Multiprocessor Real-Time Scheduling on General Purpose Operating Systems

Bridging the gap between Theory and Practice

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Abstract

From the perspective of an user the differences between desktop, server and mobile computing systems are less and less noticeable. They all ship multicore or multiprocessor CPUs and they all run similar applications on top of a General Purpose Operating Systems (GPOS). Moreover, applications runtime requirements often fall under the domain of Real-Time systems, for which not only the functional correctness of a computation, but also its timeliness is important. One of the key roles of a GPOS is then to provide, under a common interface, low-level mechanisms enabling the development of applications that can meet the needs of end users. At the same time, such an Operating System has to remain backward compatible with older platform (e.g., small singleprocessors), while being efficient throughout the whole spectrum of computing systems. On the other side, real-time academic literature usually abstracts from these requirements, as is constantly anticipating problems of future generation platforms. Unfortunately, the strong focus on theoretical aspects induces a gap between the two worlds, where knowledge created by the latter is usually hard to be applied to the former, as practical issues often arise.

This thesis wants to make the gap a bit more shallow, by developing strategies to enable the use of classical multiprocessor real-time scheduling mechanisms on a modern GPOS. To this end, we focus on the Linux kernel, as its use is nowadays widespread on all the aforementioned platforms. We first present a sofware addition to the Linux scheduler implementing Earliest Deadline First (EDF) and Constant Bandwidth Server (CBS) algorithms, going by the name of SCHED_DEADLINE; indeed, a major outcome of this work is the inclusion of SCHED_DEADLINE in the mainline version of the Linux kernel. We also detail software solutions that allow efficient real-time scheduling on a big multiprocessor system, and how the development of such solutions was eased by means of a user-space scheduler simulator. Secondly, building on top of our implementation, we report about a comparison between two classical real-time scheduling algorithms (Rate Monotonic (RM) and EDF) in order to help software developers choosing the right algorithm to schedule real-time tasks. In this part we both consider runtime overheads due to implementation choices and cache-related delays originating from the presence of memory hierachies. We finally add to the picture problems that arise when concurrent

real-time tasks share resources, and detail about theoretical extensions and a practical evaluation of one resource sharing algorithm called Multiprocessor Bandwidth Inheritance (M-BWI).

To my family and to Grazia,
you guys rock!

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Introduction

The goal of this thesis is to reduce the gap between real-time literature and industry, in the context of *General Purpose Operating Systems* (GPOSes), by developing real-time scheduling algorithms and verifying their performance on *Symmetric Multiprocessing* (SMP) systems.

This work is motivated by the widespread adoption of GPOSes on modern multiprocessors platforms to perform activities that requires certain degrees of timeliness (e.g., video processing, audio/video streaming, VoIP, etc.). The GPOS of reference is Linux, as it is nowadays widely adopted on the whole spectrum of computing platforms, ranging from small hand-held devices to cloud computing infrastructures. Even if Linux is born as a traditional GPOS, in the last years, there has been a considerable interest in using it also for real-time and control, from both academy and industry [**Edg13**]. We believe that this is mainly due to the free availability of its source code, the support for a great number of architectures and the existence of countless applications running on it. However, Linux has not been in origin tought as a *Real-Time Operating System* (RTOS), thus it lacks of mechanisms that allow a classical real-time feasibility study of the system under development; i.e., developers cannot be sure that timing requirements will be met even correctly knowing the runtime timing behavior of their applications. Indeed, POSIX-compliant fixed-priority scheduling policies, already offered by Linux, do not fit real-time users needs, as they are not much sophisticated.

Similarly to our approach, modification to the Linux kernel have been proposed, such as RT-Linux [1], proposed by Yodaiken et al. [**YB97**], and RTAI [2], proposed by Dozio et al. [**DM03**], in order to enable hard real-time computing in a Linux-like environment. In these solutions, a real-time micro-kernel layer is added between the real hardware and the Linux OS, which runs as the background/idle activity whenever there are no hard real-time tasks active in the system. This allows for respecting the very tight timing constraints (microsecond-level) typical of industrial automation and robotic applications. However, applications that want

[1]Originally supplied by FSMLabs (http://www.fsmlabs.com/), acquired by Wind River in 2007 and discontinued in 2011.

[2]https://www.rtai.org

to use real-time facilities are typically required to be adapted or rewritten. Moreover, we believe that the fact that these solutions remain bounded to the efforts of some academic entity limit their usage from industry, for which discontinuous updates to the last Linux version could represent a problem. It must be also noted the fact that a key feature like the temporal isolation property [**BLAC06**] is usually neglected and not implemented in both general purpose and real-time OSes. In fact, without such mechanism, a high priority task runs undisturbed until it blocks, indipendently from what considered at analysis/design phase. This can obviously jeopardize guarantees offered to other tasks and activities of the system, till the point the whole system becomes unusable. With our mechanisms application developers are guaranteed that the performance that their applications exibit in isolation (i.e., when run alone on the system) are not affected by other applications concurrently running on the system (and we give advices on how to cope with cases when perfect isolation cannot be guaranteed).

Among others, a modification to the Linux kernel purposely targeted for academic research is the $LITMUS^{RT}$ testbed [3], developed by the Real-Time System Group at University of North Carolina at Chapel Hill. The primary purpose of the testbed is to *provide a useful experimental platform for applied real-time system research*. Indeed, solutions developed on top of it can serve as a proof of concept for the engineering of the same solutions on plain Linux, even if a proper implementation on Linux is usually harder to be realized. As this project is not focused on industry, there are currently no plans to turn it into a production-quality system. Moreover, the API (i.e., interface to applications) is not stable and may change without warning between releases (nor the release schedule is fixed).

We instead targeted the inclusion of our contributions (at least the bigger part of them) in the mainline (stock) Linux kernel as another goal of this work. The aim is to transfer back knowledge to the Linux community, from which we copiously drew both as code base (the Linux kernel itself) and with requests for advice (through the Linux kernel mailing list). So our intent is different from $LITMUS^{RT}$ one, as we trade a certain less flexibility on our solutions with strict adherence to Linux design choices, and we do so because we firmly believe that this approach could foster a wider usage and understanding of real-time concepts by industry.

The main **contributions** of this thesis are:

- an efficient extension of the SCHED_DEADLINE patchset for the Linux scheduler for SMP systems;
- an evalutation of the performance of such a real-time extension for applications users;

[3] http://www.litmus-rt.org/

- the development of a user-space emulator of the Linux scheduler subsystem on a multi-core architecture and an evaluation of different solution to speed-up scheduling on such kind of systems;
- an experimental comparison of several configurations of two classical real-time scheduling algorithms on NUMA machines;
- a proof-of-concept implementation of the Multiprocessor Bandwidth Inheritance protocol on Linux.

An key outcome of the work performed while doing this thesis has also been the inclusion in the mainline Linux kernel (since Linux version 3.14) of almost all the code we built upon our research.

The reminder of this thesis is organized as follows. Chapter 1 discusses needed notation and background on real-time systems. Chapter 2 provides an overview on the current status of real-time scheduling mechanisms on General Purpose Operating Systems, and details about our modifications to one of such GPOSes, Linux. Moreover, it also discuss how a user-space emulator can ease real-time scheduling algorithms development on a multi-core platform. Chapter 3 builds upon both the developed mechanisms and the stock Linux scheduler to perform an experimental study of applicability of classical real-time theory on Non-Uniform Memory Access systems. Chapter 4 provides a proof-of-concept solution that allows to perform a feasiblity analysis of a real-time system using Linux even in presence of task accessing shared resources. Finally, Chapter 5 concludes with a summary of the work presented in this thesis and with the discussion of how work could extend the results presented.

CHAPTER 1

Real-Time Systems

In this chapter we provide a general introduction to real-time systems. Notation and definition are given that put the basis for the following chapters. We also detail about differences and peculiarities of Uniprocessor and Multiprocessors systems. We conclude the chapter with a taxonomy of different approaches in designing Operating Systems, and the predictability issues that may arise when such Operating Systems have to provide support for real-time applications.

1.1. Real-Time Task Model

Real-Time systems are computing systems that contain concurrent computational activities for which, not only correctness of results, but also timeliness is crucial. These computational activities are called *tasks* and each task may spawn a (potentially infinite) sequence of *jobs* during its lifetime. A *job* is thus a sequential unit of work. Timeliness requirements pertain to tasks and are embodied by *deadlines*, which represent the time before which a process should complete its execution.

We assume (unless otherwise specified, like in Chapter 4) that tasks adhere to the classical *sporadic task model* [**But11**]. The system is comprised of n real-time tasks $\tau_1, \tau_2, \ldots, \tau_n$, that constitute a *taskset* $\tau = \tau_1, \tau_2, \ldots, \tau_n$. Each task generates a sequence of jobs $\tau_{i,1}, \tau_{i,2}, \tau_{i,3}, /dots$. A job $\tau_{i,j}$ is characterized by several parameters:

Release/Arrival time ($r_{i,k}$): is the instant of time at which the job becomes ready for execution (since it has been activated by some event or condition).

Computation time ($c_{i,k}$): is the time necessary to the processor to execute the job without interruption.

Start time ($s_{i,k}$): is the time at which the job starts its execution for the first time (i.e., the processor is assigned to $\tau_{i,k}$ for the first time).

Finishing/Completion time ($f_{i,k}$): is the time at which the job completes its execution.

Relative deadline ($D_{i,k}$): is the interval of time within which the job execution should complete with respect to its release time. We usually have a single value for

every job of a certain task that is still called *relative deadline* and it is denoted as D_i (as it refers to task τ_i).

Absolute deadline ($d_{i,k}$): is the absolute instant of time by which job $\tau_{i,k}$ should complete, in order to preserve the timeliness properties of the system. The absolute deadline is computed based on the relative deadline and the release time: $d_{i,k} = r_{i,k} + D_{i,k}$.

Response time ($R_{i,k}$): is the difference between the finishing time and the release time: $R_{i,k} = f_{i,k} - r_{i,k}$.

Lateness ($L_{i,k}$): is the delay of a job completion with respect to its deadline: $L_{i,k} = f_{i,k} - d_{i,k}$. Note that, if the job completes before its deadline, its lateness is negative. Instead, when a job completes after its deadline, its lateness is posive, and in this case we say that a *deadline miss* event occurred.

Tardiness or Exceeding time ($E_{i,k}$): is the time a job stays active after its deadline: $E_{i,k} = max(0, L_{i,k})$.

Worst case execution time (C_i): shortened with WCET, is the worst (i.e., maximum) computation time of all jobs of task τ_i: $C_i = max(c_{i,k})$.

Some of the above parameters are illustrated in Figure 1.1. Usually, new arrivals are represented with upward arrows and deadlines with downward arrows.

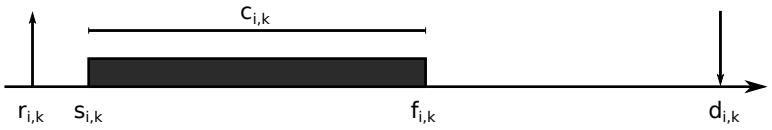


FIGURE 1.1. Typical parameters of a real-time task.

Task Periodicity Another timing characteristic that can be associated to a real-time task is the regularity of its activations. Commonly, tasks are distinguished between *periodic* and *aperiodic*.

Periodic tasks are activated (released) at regular intervals of time. Specifically, a task τ_i is said to be periodic if, for every pair of consecutive jobs, $r_{i,k+1} = r_{i,k} + T_i$, where T_i is the task *period*. The period is also used to calculate the share of system processor resource a task uses once it is activated. This parameter is called task *utilization* and corresponds to $U_i = \frac{C_i}{T_i}$. Therefore, a taskset composed by n tasks will have a total utilization of:

$$U = \sum_{i=1}^{n} U_i.$$

When activations are not strictly periodic (e.g., a task that is activated only when an aperiodic event occurs), but still there is some bounded separation among them, tasks are said to be *sporadic*. In this case, the parameter T_i denotes the

minimum separation between successive jobs of the same task, and is also called *minimum interarrival* time. We thus have: $r_{i,k+1} \geq r_{i,k} + T_i$ (note the *greater or equal* relation between release times). The sporadic task model is a generalization of the periodic task model.

Lastly, real-time tasks for which no constrain on a certain separation between consecutive activations can be given are called *aperiodic*. The only obvious observation that can be made is that jobs activations still follow a sequence, i.e. $r_{i,k+1} \geq r_{i,k}$.

1.2. Hard and Soft Real-Time Requirements

As already stated, one characteristic of real-time tasks is that for them timeliness is of key importance : a real-time job should complete (and produce a result) before its absolute deadline, otherwise the produced result, even if correct, may be too late to be useful. Depending on the criticality of this timing requirement, we can split real-time tasks in two classes: *hard* and *soft* tasks.

A task τ_i is a *hard real-time (HRT) task* if no job deadline must be missed (i.e., $E_{i,k} = 0$). *HRT systems* are comprised of HRT tasks only. A task τ_i is a *soft real-time task (SRT)* if some missed deadline are allowed for it. Systems that contain one or more SRT tasks are called *soft real-time systems*.

Hence, for HRT tasks correctness implies both correct results and no deadline misses. Instead, for SRT tasks, this notion has no single definition, being the extent of permissible deadline violations (tardiness) very application-dependent. In this thesis we adopt the notion of *bounded tardiness* [**DA08**] (i.e., each job is allowed to complete within some bounded amount of time after its deadline). It must be noted that the HRT correctness is a special case of the SRT bounded tardiness. In fact, for HRT tasks the relation $E_{i,k} = 0$ must always hold.

1.3. Real-Time Scheduling

In an Operating System (OS) kernel, the *scheduler* is responsible for choosing which task (also called process or thread) executes on each processor at any given time. In real-time systems this is done by first assigning priorities to tasks, then a real-time scheduling algorithm, implemented by the scheduler, uses such priorities to perform scheduling decisions. Furthermore, priorities can be fixed for the whole lifetime of a task, or change dynamically, according to some logic or external/internal event.

Fixed priority scheduling. In *fixed priority* scheduling, once a priority has been assigned to a task, it remains the same for the task's lifetime, and it also corresponds to the priority of every job of the task. The most popular fixed priority scheme is

the *Rate Monotonic* (RM) algorithm, for which each task τ_i gets a priority p_i that is inversely proportional to its period ($p_i \propto \frac{1}{T_i}$).

Dynamic priority scheduling. In *dynamic priority* scheduling, the priority of a task can change over time. In this thesis we focus on *Job-Level Dynamic Priority* (JLDP) algorithms, for which priority of different jobs of the same task can change, but once a priority has been assigned to a job, it remains the same till the job completion. The most widely known JLDP real-time scheduling algorithm is *Earliest Deadline First* (EDF), for which the job with the earliest absolute deadline d_i is assigned the highest dynamic priority.

Feasibility analysis. A *schedule* is the assignment (produced by the scheduler) of all jobs in the system on the available processors. Furthermore, in a valid schedule: **(i)** every processor is assigned to at most one job at any time, **(ii)** every job is scheduled on at most one processor at any time, and **(iii)** jobs are not scheduled before their release time.

A taskset τ is *feasible* on a given hardware platform if there exists a schedule (*feasible schedule*) in which every job of τ meets its deadline. A HRT taskset τ is said to be *(HRT) schedulable* on a hardware platform by algorithm A if A always produces a feasible schedule for τ (i.e., no job of τ misses its deadline under A). Moreover, A is an *optimal* scheduling algorithm if A correctly schedules every feasible task system. Relaxing the correctness notion, a SRT taskset τ is *(SRT) schedulable* under the scheduling algorithm A if the maximum tardiness is bounded.

Scheduling algorithms performance are usually compared through the *schedulable utilization bound* (or simply *utilization bound*). If $U_b(A)$ is a utilization bound for the scheduling algorithm A, then A can correctly schedule every task system with $U(\tau) \leq U_b(A)$. Note that, unless an optimal utilization bound is known for A (e.g., EDF on UP), the previous condition is only sufficient, but not necessary. In fact, there may exist a taskset τ with $U(\tau) > U_b(A)$ that is schedulable using A (e.g., RM with few tasks).

1.3.1. Uniprocessor Scheduling

In [**LL73**], Liu and Layland showed that RM is optimal among fixed-priority algorithms and they derived an utilization bound for RM for periodic task systems: $U_b = n(2^{1/n} - 1)$, that for high values of n tends to $U_b = \ln 2 \simeq 0.69$. The feasibility analysis of the RM algorithm can also be performed using a different approach called *Hyperbolic Bound* [**BBB03**]:

$$\prod_{i=1}^{n}(U_i + 1) \leq 2.$$

The test has the same complexity of the original Liu and Layland bound, but it is less pessimistic.

The EDF scheduling algorithm can schedule every feasible task system on a single-processor (UP) platform (i.e., EDF is optimal on uniprocessor systems). In fact, a taskset τ is schedulable under EDF on a uniprocessor platform if $U(\tau) \leq 1 = U_b(EDF)$ [**LL73**].

1.3.2. Multiprocessor Scheduling

Two basic approaches exist for scheduling real-time tasks on multiprocessor systems. In the *partitioned* approach, each task is statically assigned to a single processor and migration between processors is not allowed; in the *global* approach, tasks can freely migrate and execute on any processor.

Partitioned scheduling algorithms have the advantage that uniprocessor scheduling algorithms, and feasibility tests, can be separately used on each processor. Contrariwise, they also require to solve a bin-packing-like problem to assign tasks to processor. Being similar to a bin-packing problem (NP-hard in the strong sense), the assignment of tasks to processors is usually performed using heuristics on tasks utilizations (e.g., first-fit, best-fit, next-fit, worst-fit).

Under global approaches, tasks are conceptually selected from a single run-queue (see how this is implemented in the Linux kernel below) and may migrate between processors. At any instant of time, at most M (on an platform composed of M processors) ready jobs with the highest priority execute on the M processors. Focusing only on the EDF algorithm, we see that, similarly to UP EDF, in a HRT system the *global* EDF (G-EDF) scheduling algorithm also requires up to $(2 \cdot U(\tau) - 1)$ processors to feasibly schedule a taskset τ where the maximum per-task utilization is $max(u_i) \leq 1/2$. However, in 1978 Dhall and Liu noted that on multiprocessor platforms there exist task sets with total utilization close to 1.0 that cannot be HRT scheduled by G-EDF or *global* RM (G-RM) [**DL78**]. Mainly because of this fact, global approaches are not usually adopted for HRT systems. Nonetheless, when SRT systems are considered, G-EDF ensure bounded tardiness as long as the system is not overutilized [**DA08**].

As a compromise between the two approaches, *clustered* scheduling has been proposed [**CAB07**], that aims to alleviate limitations of partitioned and global scheduling on large multicore platforms. Under clustered algorithm, the platform is partitioned into *clusters* of cores that share a cache and tasks are statically assigned to clusters, but are globally scheduled within each cluster.

1.4. Real-Time Operating Systems

This section provides an overview of the most important issues, regarding predictability and design opportunities, that can be faced when implementing Real-Time Operating Systems (RTOS). The main focus is (here and in the rest of the

thesis) on Linux-based RTOS, as Linux is nowadays largely used by both academia and industry to perform activities that fall in the real-time domain.

1.4.1. Predictability Issues

Linux has been designed as a General Purpose Operating Systems, so it is no wonder that it can experiment problems with HRT. In particular, the main issues of predictability under Linux are due to:

- Non-preemptable critical sections. Several execution paths in the kernel cannot be preempted and interrupts are disabled during the execution of certain IRQ managements routines. These factors can cause priority inversions and thus unpredictability for real-time activities.
- Non-predictable duration of IRQ management routines. Even though Linux adopts a split-interrupt management schema, the duration of IRQ managements routines is non predictable and can thus affect predictability of the system.
- Throughput oriented scheduling. Linux has been designed to be throughput oriented. Scheduling decisions on multiprocessor systems tend to evenly distribute the load among available processors, without considering priorities or any effect related to the presence of caches and memories. This may cause unneeded migrations, high overhead and tasks execution time variance, impacting system predictability.

In order to solve these problems (in what follows we will explicitly deal with the last point) several approaches to modify Linux has been proposed, that can be divided into *mono-* and *dual-kernel* variants.

Mono-kernel approach

Under this approach, high predictability is achieved addressing the aforementioned limitation with modification of the Linux kernel. This approach is in common between some commercial RTOSs (e.g., MontaVista Linux [1], timesys [2], etc.) and by the open-source PREEMPT_RT patch for the Linux kernel [3]. This patch allows nearly all of the kernel to be preempted, with the exception of a few small region of code. This is achieved by replacing most kernel spinlocks (a simple single-holder lock) with mutexes that support priority inheritance, as well as moving all interrupts and software interrupts to kernel threads. It has to be noted that our contributions are orthogonal to this first approach (user of the last version of the PREEMPT_RT patch, based on Linux version 3.14, can find our contributions already included) and can actually address point three above (whereas PREEMPT_RT solves point one and two).

[1] http://www.mvista.com/solution-real-time.html

[2] http://www.timesys.com/

[3] https://rt.wiki.kernel.org/index.php/CONFIG_PREEMPT_RT_Patch

Dual-kernel approach

Under this approach, a virtualization layer is employed to concurrently execute two (or more) operating systems on the same hardware. Although this additional layer may introduce additional overhead and latencies, this second approach allows to isolate GPOSs like Linux for RTOSs, thus avoiding the predictability issues detailed above. Linux-derived RTOSs like RTAI [4] and RTLinux [5] employ a dual-kernel approach to meet the requirements of HRT tasks. Dual-kernel approaches are also employed by commercial RTOSs like VxWorks [6].

1.4.2. Design Opportunities

Free availability of the Linux kernel source code gives the impression that the desing space for modifications of its internal meachanisms is boundless. And this is almost true, unless one doesn't plan to propose such modifications to the Linux kernel development community. The community is responsible, through the work of subsystems maintainers, for guaranteeing that Linux remains efficient on a vast set of hardware platform. This is accomplished via a thorough review process of patches proposed to the community, and, understandably, big changes and complete rewrites of core subsystems are generally frowned upon. The Linux scheduler makes no exception; being one of the most important subsystems, it is actually very refractory to modifications.

The main design characteristic that constrain design space for new modifications are:

- Low overhead. The scheduler must be extremely fast in deciding which process to run next. No heavy operations are thus allowed. Purpose built data structures are usually employed to speed up scheduling decisions.
- Seamless integration in the scheduler framework. The Linux scheduler has a modular design (see Section 2.1.2. New scheduling policies must fit in this desing.
- Distributed design. As we will describe in more details in Section 2.1.2, the Linux scheduler performs scheduling decision in a distributed manner. In particular, each CPU has its own runqueue and global scheduling is achieved through tasks migrations among runqueues. There is flexibility in deciding how migration decisions happen.
- Fast acceptance/refusal of new tasks entering the system. The scheduler is also responsible for deciding if a new task can enter the system (e.g., if the user has enough permissions for choosing a scheduling policy). Such

[4]https://www.rtai.org

[5]Originally supplied by FSMLabs (http://www.fsmlabs.com/), acquired by Wind River in 2007 and discontinued in 2011.

[6]http://www.windriver.com/products/vxworks/index.html

decisions must be quickly perfomed. A heavy machinery, even if more accurate, is usually not allowed.

We decided to adhere to these constraints as we planned inclusion of our modifications in the mainline Linux kernel right from the beginning.

CHAPTER 2

Real-Time Scheduling on General Purpose Operating Systems

2.1. Introduction

In this and in the following chapter, we argue that real-time scheduling is indeed possible also on top of General Purpose Operating Systems (GPOS). We will focus on the Linux kernel, given its widespread adoption and the availability of its source code, but the obtained results can be generalized to other GPOS having the same structure of Linux and running on top of nowadays multi-core and multiprocessor machines. We specifically address issues detailed in Section 1.4 giving an overview of the design choices we made considering the design constraints of Section 1.4.2. Furthermore, we keep our focus close to the point of view of an application developer. In effect, we advise the use of the following technologies as a way to improve efficiency and predictability of a vast number of classes of applications.

2.1.1. Contributions

In this chapter, we first present the **implementation of a global EDF scheduler** in Linux, called `SCHED_DEADLINE` [**FCTS09**]. We also show how we optimized global scheduling on SMP systems using a heap data structure. After describing the base real-time scheduler of Linux (Section 2.1.2), and our implementation (Section 2.2), we compare its performance against the global POSIX-compliant fixed priority scheduler of Linux and with a previous version of `SCHED_DEADLINE` (Section 2.2.7). The results show that **using appropriate data structures it is indeed possible to build efficient and scalable global real-time schedulers**. SCHED_DEADLINE is part of the mainline Linux kernel since version 3.14.

We then propose PRAcTISE (PeRformance Analysis and TestIng of real-time multicore SchEdulers) for the Linux kernel: it is a **framework for developing, testing and debugging scheduling algorithms in user space** before implementing them in the Linux kernel. In addition, PRAcTISE allows to compare different implementations by providing early estimations of their relative performance. In this way, the most appropriate data structures and scheduler structure can be chosen and evaluated in user-space. Compared to other similar tools, like

LinSched, the proposed framework allows **true parallelism thus permitting a full test in a realistic scenario**.

2.1.2. The Linux Scheduler

Since release 2.6.23, the Linux scheduler is implemented as a modular framework that can be easily extended. The current structure has been implemented by Ingo Molnar as a replacement of the previous O(1) scheduler. The structure consists of a core block, providing basic funtionalities, and a set of *scheduling classes*, each encapsulating one or more specific *scheduling policy*. Scheduling policies determine when and how tasks are selected to run. Figure 2.1 shows the set of scheduling policies traditionally available in Linux (i.e., until version 3.13).

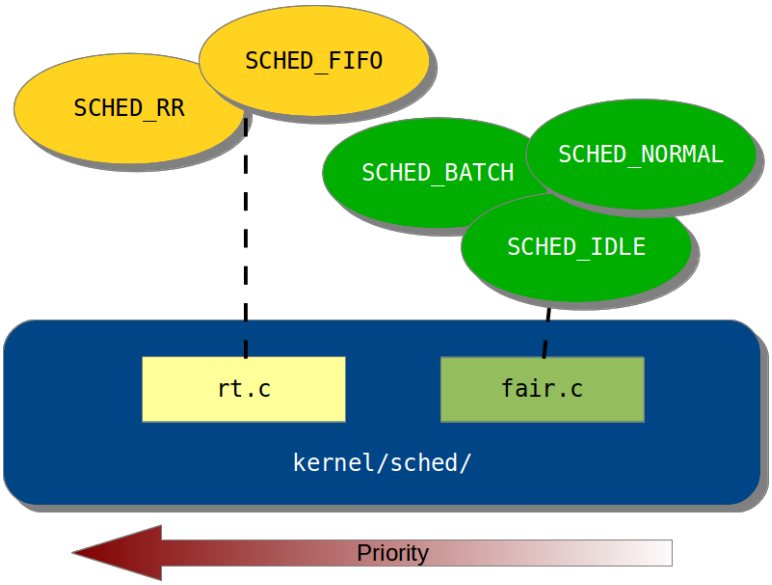


FIGURE 2.1. Linux modular scheduling frameworki (until version 3.13).

Each scheduling policy of Figure 2.1 belongs to one of the two scheduling classes Linux had before version 3.13. First scheduling class, implemented in `kernel/sched/fair.c`, is intended for fair scheduling of non-real-time activities (`SCHED_NORMAL`, `SCHED_BATCH` and `SCHED_IDLE` policies). Second scheduling class is providing fixed priority real-time scheduling (`SCHED_FIFO` or `SCHED_RR` policies), following the POSIX 1001.3b [**IEE04**] specification and is implemented in `kernel/sched/rt.c`. For what concern tasks priorities, scheduling classes are stacked upon each other, where lower classes have also lower priorities. As it is indicated in the picture, real-time scheduling policies have higher priority than fair scheduling policies, i.e., tasks belonging to the latter are always preempted by tasks scheduled by the former. In this thesis we will focus specifically on real-time scheduling policies.

Run-queues, masks and locks

To keep track of active tasks, the scheduler uses a data structure called *runqueue*. There is one runqueue for each CPU and they are managed separately in a distributed manner. Every runqueue is protected by a spin-lock to guarantee correctness on concurrent updates. Runqueues are modular, in the sense that there is a separate sub-runqueue for each scheduling class. Tasks are *enqueued* on some runqueue when they wake up and are *dequeued* when they are suspended.

Key components of the fixed priority sub-runqueue are:

- a priority array on which tasks are actually queued;
- fields used for load balancing;
- fields to speed up decisions on a multiprocessor environment.

The fixed priority scheduling class already supports global scheduling. Tasks are migrated across CPUs (runqueues) following an active load balancing approach that is realized through *push* and *pull* operations, see below.

An additional data structure, called `cpupri`, is used to reduce the amount of work needed for a push operation. This structure tracks the priority of the highest priority task in each runqueue. The system maintains the state of each CPU with a 2 dimensional bitmap: the first dimension is for priority class and the second for CPUs in that class. Therefore a push operation can find a suitable CPU where to send a task in $O(1)$ time, since it has to perform a two bits search only (if we don't consider affinity restriction).

Push and pull operations

When a task is activated on CPU k, first the scheduler checks the local runqueue to see if the task has higher priority than the executing one. In this case, a preemption happens, and the preempted task is inserted at the head of the queue; otherwise the waken-up task is inserted in the proper runqueue, depending on the state of the system. In case the head of the queue is modified, a *push* operation is executed to see if some task can be moved to another queue. When a task suspends itself (due to blocking or sleeping) or lowers its priority on CPU k, the scheduler performs a *pull* operation: it looks at the other run-queues to see if some other higher priority tasks need to be migrated to the current CPU. Pushing or pulling a task entails modifying the state of the source and destination runqueues: the scheduler has to dequeue the task from the source and then enqueue it on the destination runqueues.

2.2. SCHED_DEADLINE

At the time of writing, a new scheduling class/policy has been merged in the Linux kernel, called `SCHED_DEADLINE` [**FCTS09**]. It implements partitioned, clustered and global EDF scheduling with hard and soft reservations [1]. `SCHED_DEADLINE` is seamless integrated in the Linux scheduler modular framework. The scheduling policy is implemented in `kernel/sched/deadline.c` and extends the set of traditionally available scheduling policies as can be seen in Figure 2.2.

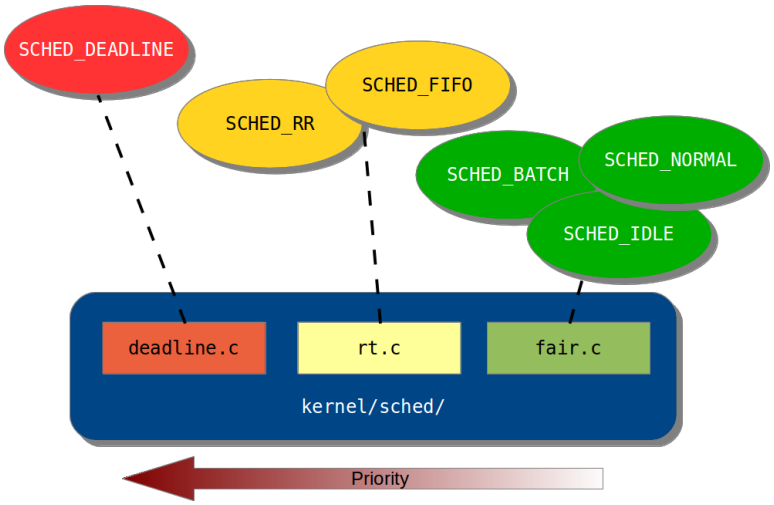


FIGURE 2.2. Linux modular scheduling framework, since Linux 3.14.

In this section [2], we describe the implementation of SCHED_DEADLINE and we evaluate its properties running synthetic benchmarks and real applications on a Linux systems. We then detail about a heap data structure we designed and developed to optimize access to the earliest deadline tasks (the implementation is now part of `SCHED_DEADLINE` [**FCTS09**]). After describing the base real-time scheduler of Linux (Section 2.1.2), and our implementation (Section 2.2), we compare its performance against the global POSIX-compliant fixed priority scheduler of Linux and with a previous version of `SCHED_DEADLINE` (Section 2.2.7). The results show that **using appropriate data structures it is indeed possible to build efficient and scalable global real-time schedulers**.

The code developed during the experimental evaluation phase of this part can be downloaded by following the instructions on this page: `http://retis.sssup.it/~jlelli/sched-deadline.php`.

[1] Full source code merged in mainline since Linux 3.14.

[2] Claudio Scordino contributed to Section 2.2.2 and Luca Abeni performed experiments of Sections 2.2.3, the content of these Sections has been submitted for publication.

```
struct dl_rq {
    struct rb_root rb_root;
    struct rb_node *rb_leftmost;
    unsigned long dl_nr_running;

#ifdef CONFIG_SMP
    struct {
        /* two earliest tasks in queue */
        u64 curr;
        u64 next; /* next earliest */
    } earliest_dl;
    int overloaded;
    unsigned long dl_nr_migratory;
    unsigned long dl_nr_total;
    struct rb_root pushable_tasks_root;
    struct rb_node *pushable_tasks_leftmost;
#endif /* CONFIG_SMP */
};
```

FIGURE 2.3. `struct dl_rq` extended

2.2.1. Implementation Details

The approach used for the implementation is the same used in the Linux kernel for the fixed-priority scheduler. This is usually called *distributed run-queue*, meaning that each CPU maintains a private data structure implementing its own ready queue and, if global scheduling is to be achieved, tasks are migrated among processors when needed.

In more details:

- the tasks of each CPU are kept into a CPU-specific run-queue, implemented as a *red-black tree* ordered by absolute deadlines;
- tasks are migrated among run-queues of different CPUs for the purpose of fulfilling the following constraints:
  - on m CPUs, the m earliest deadline ready tasks run;
  - the CPU affinity settings of all the tasks is respected.

Migration points are the same as in the fixed priority scheduling class. Decisions related to push and pull logic are taken considering deadlines (instead of priorities) and according to tasks affinity and system topology. The data structure used to represent the EDF ready queue of each processor has been modified, as shown in Figure 2.3 (new fields are the one inside the `#ifdef CONFIG_SMP` block).

- `earliest_dl` is a per-runqueue data structure used for "caching" the deadlines of the first two ready tasks, so to facilitate migration-related decisions;
- `dl_nr_migratory` and `dl_nr_total` represent the number of queued tasks that can migrate and the total number of queued tasks, respectively;
- `overloaded` serves as a flag, and it is set when the queue contains more than one task;
- `pushable_tasks_root` is the root of the red-black tree of tasks that can be migrated, since they are queued but not running, and it is ordered by increasing deadline;
- `pushable_tasks_leftmost` is a pointer to the node of `pushable_tasks_root` containing the task with the earliest deadline.

A *push* operation tries to move the first ready and not running task of an overloaded queue to a CPU where it can execute. The best CPU where to push a task is the one which is running the task with the latest deadline among the m executing tasks, considering also the constraints due to the CPU affinity settings. A *pull* operation tries to move the most urgent ready and not running tasks among all tasks on all overloaded queues in the current CPU.

2.2.2. User-level API

The existing system calls `sched_setscheduler()` and `sched_getscheduler()` have not been extended, due to the binary compatibility issues that modifying the `sched_param` data structure would have raised for existing applications. Therefore, two new system calls called `sched_setattr()` and `sched_getattr()` have been introduced. These syscalls also support the other existing scheduling policies — i.e., the interpretation of the arguments depends on the selected policy. Therefore, they are expected to replace the previous system calls (which will be left to not break existing applications). The prototype of these new system calls is shown in Figure 2.4.

2.2.3. Experiments

Greedy tasks

As a first experiment, we have used SCHED_DEADLINE to schedule one periodic task (that executes for 1ms every 4ms) and two greedy tasks (tasks which never blocks, and try to consume all the CPU time) scheduled by two CBSs (1ms, 6ms) and (1ms, 10ms). Figure 2.5 shows a segment of the schedule. As periodic task is the one with most strict timing requirements (relative deadline is 4ms) it gets always scheduled when it wakes up. Greedy 1 is also higher priority than Greedy 2, and it preempts the latter during the first activation (remember that priorities are dynamic, the behavior is thus relative to this particular timing window). Another thing to notice is that greedy tasks are throttled once they try to execute for more

```
#include <sched.h>

struct sched_attr {
    u32 size;
    u32 sched_policy;
    u64 sched_flags;

    /* SCHED_OTHER, SCHED_BATCH */
    s32 sched_nice;

    /* SCHED_FIFO, SCHED_RR */
    u32 sched_priority;

    /* SCHED_DEADLINE */
    u64 sched_runtime;
    u64 sched_deadline;
    u64 sched_period;
};

int sched_setattr(pid_t pid,
    const struct sched_attr *attr);

int sched_getattr(pid_t pid,
    const struct sched_attr *attr,
    unsigned int size);
```

FIGURE 2.4. `SCHED_DEADLINE` API

than the allowed budget (red lines in the figure), while the periodic task always goes to sleep before exausting its budget, and it is never throttled.

This experiment shows that SCHED_DEADLINE is capable of creating an effective isolation between the running tasks, so that greedy, buggy or misbehaving tasks cannot affect the execution of the other running tasks.

Synthetic Real-Time Workloads

In order to show how SCHED_DEADLINE allows to properly schedule real-time applications, some sets of periodic real-time tasks have been randomly generated with taskgen [**ESD10**] and executed by a user-level application (named `rt-app`) either under the SCHED_DEADLINE or the `SCHED_OTHER` (i.e., CFS) scheduling

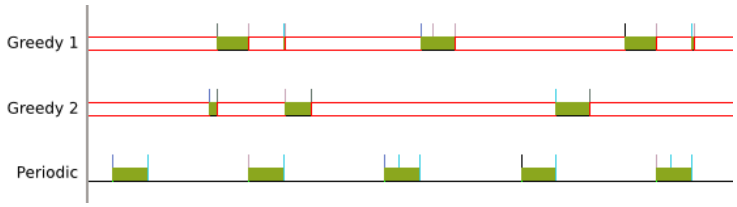


FIGURE 2.5. SCHED_DEADLINE serving a periodic task and two CPU hungry (greedy) tasks.

TABLE 2.1. Percentage of missed deadline for partitioned scheduling, as a function of the load $U = \sum \frac{C_i}{P_i}$ expressed as a percentage.

$U(\%)$	SCHED_DEADLINE	SCHED_OTHER
60%	0%	0.58%
70%	0%	0.9%
80%	0%	2.61%
90%	0%	5.88%

policies. When using SCHED_DEADLINE , each task τ_i has been assigned a runtime Q_i slightly larger than its execution time C_i, and a server period T_i equal to the task period P_i.

This experiment has been performed considering both partitioned scheduling and global scheduling. In the partitioned scheduling case, tasks were statically bound to a CPU core (using the Linux `cpuset` mechanism), and the load on each core increased from 0.6 (60%) to 0.9 (90%). Notice that the $U = 1$ case (100% CPU utilization) has been avoided in order to leave some spare time for the other tasks running in the system, so that the OS is not starved by SCHED_DEADLINE tasks. For each CPU load, 50 tasksets were randomly generated.

Table 2.1 shows the percentage of missed deadlines when using SCHED_DEADLINE or CFS to schedule the tasksets. As it can be noticed from the table, SCHED_DEADLINE is able to avoid any missed deadline (because the load on each core is smaller than 1, and because the CBS parameters have been assigned in order to exploit the so-called *hard schedulability* property of the CBS). On the other hand, CFS performs pretty well, but is not able to avoid missing deadlines.

After testing partitioned scheduling, the experiment has been repeated by configuring SCHED_DEADLINE to do global EDF scheduling. In this case, the Linux `cpuset` mechanism is not used and tasks are able to migrate between all of the available CPU cores. Since these experiments have been performed using 4 of the 6 cores provided by the Xeon CPU, the tasksets have been generated with a total

TABLE 2.2. Percentage of missed deadline for global scheduling, as a function of the load $U = \sum \frac{C_i}{P_i}$.

$U(\%)$	SCHED_DEADLINE	SCHED_OTHER
340%	0.77%	3.75%
350%	0.78%	6.17%
360%	1.29%	6.92%
370%	1.69%	8.52%
380%	2.38%	10.62%
390%	3.54%	14.15%

utilization ranging from $U = 3.4$ (340%) to $U = 3.9$ (390%). Again, the $U = 4$ case has not been considered in order to avoid starving the system. As for the previous experiment, 50 taskset per CPU load have been randomly generated.

Table 2.2 shows the percentage of missed deadlines when using SCHED_DEAD LINE or CFS to schedule the tasksets: in case of global scheduling, EDF is not able to guarantee that no deadline will be missed, so SCHED_DEADLINE experiences some missed deadlines. The standard CFS scheduler, however, exhibits a percentage of missed deadlines that is more than 4 times the percentage experienced by the CBS.

These sets of experiments show two things:

- On uni-processor systems or when the tasksets can be statically partitioned between multiple CPUs / CPU cores, SCHED_DEADLINE is suitable to schedule hard real-time tasks (no missed deadlines);
- On multi-processor (or multi-core) systems where the taskset cannot be statically partitioned between CPUs / cores and global scheduling must be used, SCHED_DEADLINE allows to reduce the number of missed deadlines (respect to other CPU schedulers) improving the performance of soft real-time tasks.

SCHED_DEADLINE on a Real Application

After showing how SCHED_DEADLINE helps in respecting the timing constraints of real-time tasks using some randomly generated synthetic workloads, the effectiveness of the new scheduling policy is now shown on a real application. In particular, we have performed a set of tests using MPlayer[3], a simple yet widely used and powerful video player. Being single-threaded, it can be easily scheduled through a single reservation.

[3] `http://www.mplayerhq.hu`

MPlayer has been modified to measure some important quality of service metrics when reproducing a video: the *Inter-Frame Time* (IFT) — defined as the difference between the display time of the current and the previous frame — and the *Audio / Video desynchronisation* (A/V Desynch) — defined as the difference between the Presentation TimeStamp (PTS) of the currently reproduced audio sample and the PTS of the currently reproduced video frame. Variations in the IFT can have a bad impact of the perceived video quality, because the video does not play smoothly, while large values of the A/V Desynch affect the quality of the reproduced media because audio and video do not appear synchronised (think about lip synch).

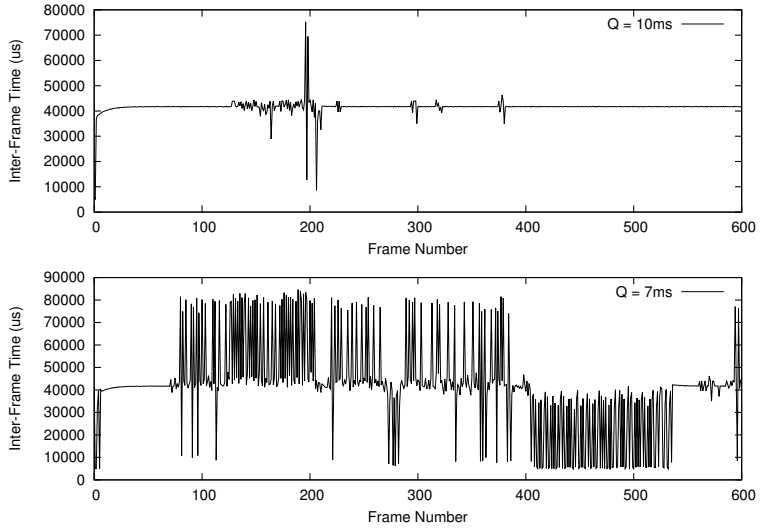


FIGURE 2.6. Inter-Frame Times for MPlayer scheduled with different SCHED_DEADLINE parameters.

MPlayer has then been used to play a HD movie (with H.264 video and AAC+ audio), and scheduled by using SCHED_DEADLINE with a period equal to the expected IFT (1 / fps) and a runtime (maximum budged) ranging from $5ms$ to $20ms$. Since the video frame rate is 23.976fps, the expected IFT is $1000000/23.976 = 41708\mu s$. As expected, if the runtime was large enough, the IFT was stable around the expected value of $41708\mu s$. Decreasing the maximum budged, some jitter started to be visible in the IFT. Finally, for small values of Q_i, the IFT was out of control. Figure 2.6 shows the IFT measured for the first 600 frames with a value of the maximum runtime near to the one needed to decode without issues ($Q_i = 10ms$) and a smaller value, which created issues and a non fluid playback ($Q_i = 7ms$).

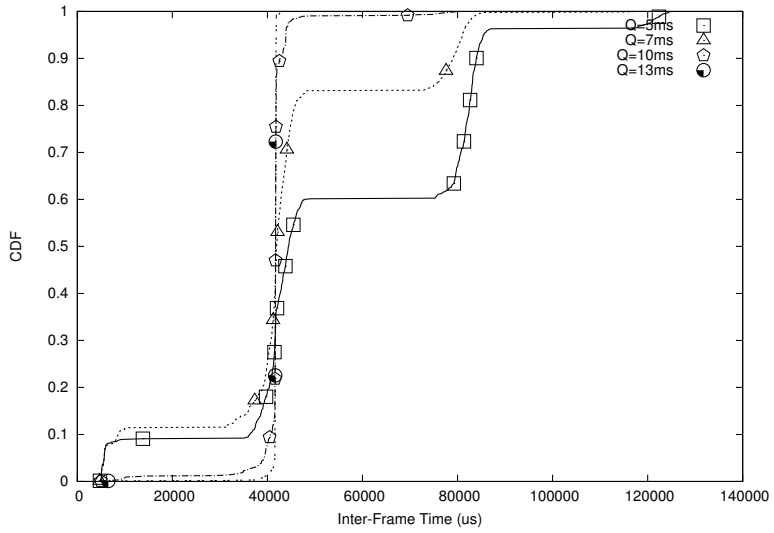


FIGURE 2.7. Cumulative Distribution Function of the MPlayer's Inter-Frame Times for different values of the SCHED_DEADLINE maximum runtime.

Figure 2.7 shows the impact of different Q_i values on the IFTs by plotting their Cumulative Distribution Function (CDF). Notice how increasing Q_i allows to make the CDF more similar to a step function (indicating that MPlayer has a probability near to 1 to play the video always at the correct rate), at the cost of dedicating more CPU time to MPlayer's execution. This experiment shows how SCHED_DEADLINE allows to respect the temporal constraints of real applications (and not only synthetic benchmarks), and to find proper trade-offs between QoS and CPU usage.

Finally, Figure 2.8 displays the A/V Desynch experienced for different values of Q_i, showing again how SCHED_DEADLINE can be used to control the quality perceived by a user and to guarantee the proper behaviour of time-sensitive applications.

Summing up, SCHED_DEADLINE provides a good amount of control over the real-time performance of real applications, because it allows to better specify the applications' parameters and requirements: since the user can communicate to the scheduler some temporal constraints to be respected (in the form of a period T_i and a runtime Q_i), the scheduler can do a better work in trying to respect these constraints.

Using SCHED_DEADLINE to Control the Application Throughput

The previous experiments showed how SCHED_DEADLINE allows to respect the deadlines of real-time applications and to properly serve "legacy" time-sensitive

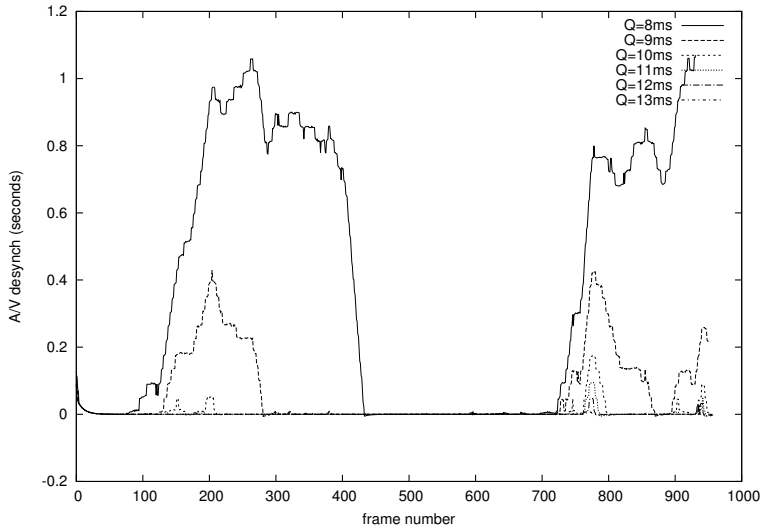


FIGURE 2.8. A/V Desynch for MPlayer's scheduled by SCHED_DEADLINE with different values of the maximum runtime.

applications that are not specifically been designed to be used with SCHED_DEAD LINE (controlling the trade-off between their quality and CPU usage). However, the temporal isolation property provided by SCHED_DEADLINE also allows to control the throughput of non real-time (or time-sensitive) applications. Hence, a last set of experiments has been performed to test this possibility.

The application used in these experiments is KVM [**KKL+07**] used to run a Virtual Router (VR) [**AKLB13**]. This VR is implemented by executing a software router (a Linux-based OS running Quagga[4] as a routing daemon and using the Linux kernel as a data plane) in a KVM virtual machine. KVM creates a user-space thread (the vCPU thread) for executing the software router and uses a kernel thread (the vhost-net kernel thread) to move packets between the virtualised software router and the physical network cards. Since both of these threads are particularly CPU-hungry (when the rate of packets to be routed is too high), SCHED_DEADLINE can be used their CPU usage, seeing how the fraction of CPU reserved to these threads can affect the router performance.

In this experiment, performed by using VRKit [**AK13**], the VR has been fed with small (64 bytes) UDP packets, while scheduling the vCPU thread and the vhost-net kernel thread with SCHED_DEADLINE and reserving different percentages of CPU time to them (remember that the percentage of CPU time reserved to

[4]`http://www.nongnu.org/quagga`

a task scheduled by SCHED_DEADLINE with runtime Q_i and period T_i is equal to $\frac{Q_i}{T_i}$ in percentage).

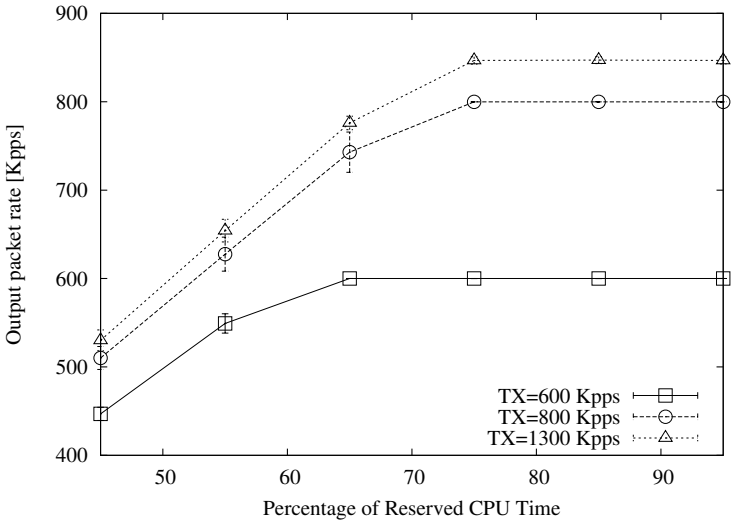


FIGURE 2.9. Throughput of a KVM-based virtual router as a function of the percentage of CPU time reserved to the vCPU thread.

Figures 2.9 and 2.10 plot the rate of routed packets as a function of Q_i/T_i (in percentage) when SCHED_DEADLINE is used to schedule the vCPU thread and the vhost-net kernel thread. The experiment has been performed for various input packet rates, but only some "interesting" lines are reported in the figures: a line corresponding to an underloaded VR (increasing Q_i/T_i, the routed packets rate reaches the input rate, and the the line becomes flat - reserving more CPU time to KVM cannot improve the performance), a line near to the overload, and a line corresponding to an overloaded VR (maximum possible input packet rate - in this case, even when the KVM thread is reserved 95% of the CPU time the routed packets rate cannot reach the input rate). The interesting thing to be noticed is that the VR performance (the routed packets rate) increases almost linearly with the ration Q_i/T_i (fraction of CPU time reserved to the vCPU or vhost-net thread). This shows that SCHED_DEADLINE allows to easily control the performance (even non real-time performance) of applications by allowing them to execute for a well-specified fraction of the CPU time.

2.2.4. Data Structures for Efficient Global Scheduling

Following sections are an extract from a paper by Lelli et al. [**LLFC11**]. As the paper focuses on different alternatives in implementing efficient global scheduling, we decided to keep here its incremental approach. The reader could be confused by

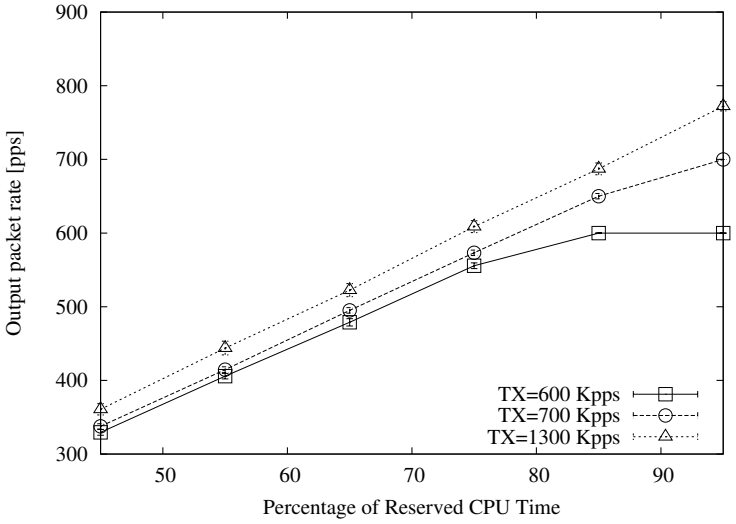


FIGURE 2.10. Throughput of a KVM-based virtual router as a function of the percentage of CPU time reserved to the vhost-net kernel thread.

the name we give to each incremental modification, though. Section 2.2.7 clarifies the naming, but we anticipate it for clearness. In the following we name:

- `original`: an old `SCHED_DEADLINE` implementation (before Linux 3.14), that didn't provide efficient global scheduling;
- `fmask`: `original` plus changes described in Section 2.2.5;
- `heap`: `original` plus the heap described in Section 2.2.6 (this corresponds almost completely to what we have today);
- the reference Linux scheduler is denoted with `SCHED_FIFO`.

Moreover, the reader should refer back to Section 2.2.2 for an introduction on the topic (basic data structures and mechanisms).

2.2.5. Idle processor improvement

The push mechanism core is realized in a small function that finds a suitable CPU for a to-be-pushed task. The operation can be easily accomplished on a small multi-core machine (for example a quad-core) just by looking at all queues in sequence. The original `SCHED_DEADLINE` implementation realizes a complete loop through all cores for every push decision (pseudo-code on Figure 2.11). The execution time of such function increases linearly with the number of cores, therefore it does not scale well to systems with large number of cores.

A simple observation is that on systems with large number of processors and relatively light load, many CPUs are idle most of the time. Therefore, when a task wakes up, there is a high probability of finding an idle CPU. To improve the

```
cpu_mask push_find_cpu(task) {
    for_each_cpu(cpu, avail_cores) {
        mask = 0;
        if (can_execute_on(task, cpu) &&
            dline_before(task, get_curr(cpu)))
            mask |= cpu;
    }
    return mask;
}
```

FIGURE 2.11. Find CPU eligible for push.

```
cpu_mask push_find_cpu(task) {
    if (dlf_mask & affinity)
       return (dlf_mask & affinity);
    mask = 0;
    for_each_cpu(cpu, avail_cores) {
        if (can_execute_on(task, cpu) &&
            dline_before(task, get_curr(cpu)))
            mask |= cpu;
    }
    return mask;
}
```

FIGURE 2.12. Using idle CPU mask.

execution time of the push function, we can use a bitmask that stores the idle CPUs with a bit equal to 1. On a 64-bit architecture, we can represent the status of up to 64 processors by using a single word. Therefore, the code of Figure 2.11 can be rewritten as in Figure 2.12, where `dlf_mask` is the mask that represents idle CPUs, and the loop is skipped (returning all suitable CPUs to the caller) if is it possible to push the task on a free CPU.

This simple data structure introduces little or no overhead for the scheduler and significantly improves performance figures in large multi-core systems (more on this later). Updates on `dlf_mask` are performed in a thread-safe way: we use a low level `set_bit()` provided in Linux which performs an atomic update of a single bit of the mask.

2.2.6. Heap Data structure

When the system load is relatively high, idle CPUs tend to be scarce. Therefore, we introduce a new data structure to speed-up the search for a destination

CPU inside a push operation. The requirements for the data structure are: $O(1)$ complexity for searching the best CPU; and less-than-linear complexity for updating the structure. The classical heap data structure fulfils such requirements as it presents $O(1)$ complexity for accessing to the first element, and $O(\log n)$ complexity for updating (if contention is not considered). Also, it can be implemented using a simple array. We developed a *max heap* to keep track of deadlines of the earliest deadline tasks currently executing on each runqueue. Deadlines are used as keys and the *heap-property* is: if B is a child node of A, then $deadline(A) \geq deadline(B)$. Therefore, the node in the root directly represent the CPU where the task need to be pushed.

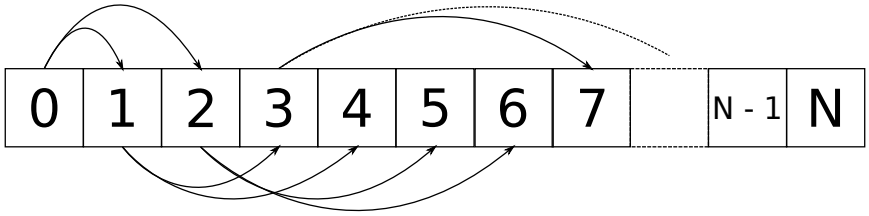


FIGURE 2.13. Heap implementation with a simple array.

A node of the heap is a simple structure that contains two fields: a deadline as key and an `int` field representing the associated CPU (we will call it `item`). The whole heap is then self-contained in another structure as described in Figure 2.14:

- `elements` contains the heap; `elements[0]` contains the root and the node in `elements[i]` has its left child in `elements[2*i]`, its right child in `elements[2*i+1]` and its parent in `elements[i/2]` (see Figure 2.13);
- `size` is the current heap size (number of non idle CPUs);
- `cpu_to_idx` is used to efficiently update the heap when runqueues state changes, since with this array we keep track of where a CPU resides in the heap;
- `free_cpus` accounts for idle CPUs in the system.

```
struct dl_heap {
    spinlock lock;
    int size;
    int cpu_to_idx[NR_CPUS];
    item elements[NR_CPUS];
    bitmask free_cpus;
};
```

FIGURE 2.14. Heap structure.

Special attention must be given to the `lock` field. Consistency of the heap must be ensured on concurrent updates: every time an update operation is performed, we force the updating task to spin, waiting for other tasks to complete their work on the heap. This kind of coarse-grained lock mechanism simplifies the implementation but it increases contention and overhead. In the future, we will look for alternative lock-free implementation strategies.

Potential points of update for the heap are enqueue and dequeue functions. If something changes at the top of a runqueue, a new task starts executing becoming the so-called *curr*, or the CPU becomes idle, the heap must be updated accordingly. We argued, and then experimented, an increase in overhead for the aforementioned operations, but we will show in Section **??** data that suggest this price is worth paying in comparison with push mechanism performance improvements.

With the introduction of the heap, code in Figure 2.12 can be changed as in Figure 2.15, where `maximum(...)` returns the heap root. As we can see from the pseudo-code we first try to push a task to idle CPUs, then we try to push it on the *latest deadline* CPU; if both operations fail, the task is not pushed away.

This kind of functioning is compliant with classical global scheduling, as it performs continuous load balancing across cores: rather than compacting all tasks on few cores we prefer every core share an (as much as possible) equal amount of real-time activities.

2.2.7. Evaluation

Experimental setup. The aim of the evaluation is to measure the performance of the new data structures compared with the reference Linux implementation

```
cpu_mask push_find_cpu(task) {
    if (dl_heap->free_cpus & affinity)
        return (dl_heap->free_cpus & affinity);
    if (maximum(dl_heap) & affinity)
        return maximum(dl_heap);
    mask = 0;
    for_each_cpu(cpu, avail_cores) {
        if (can_execute_on(task, cpu) &&
            dline_before(task, get_curr(cpu)))
            mask |= cpu;
    }
    return mask;
}
```

FIGURE 2.15. Find eligile CPU using a heap.

(SCHED_FIFO) and the original SCHED_DEADLINE implementation. Since all mechanisms described so far share the same structure (i.e. distributed runqueues, and push and pull operations for migrating tasks), we measures the average number of cycles of the main operations of the scheduler: to enqueue and dequeue a task from one of the runqueues; the push and pull operations.

We conducted our experiments on a Dell PowerEdge R815 server equipped with 64GB of RAM, and 4 AMD^R Opteron^{TM} 6168 12-core processors (running at 1.9 GHz), for a total of 48 cores. The memory is globally shared among all the cores, and the cache hierarchy is on 3 levels (see Figure 2.16), private per-core 64 KB L1D and 512 KB L2 caches, and a global 10240 KB L3 cache. The R815 server was configured with a Debian Sid distribution running a patched 2.6.36 Linux kernel.

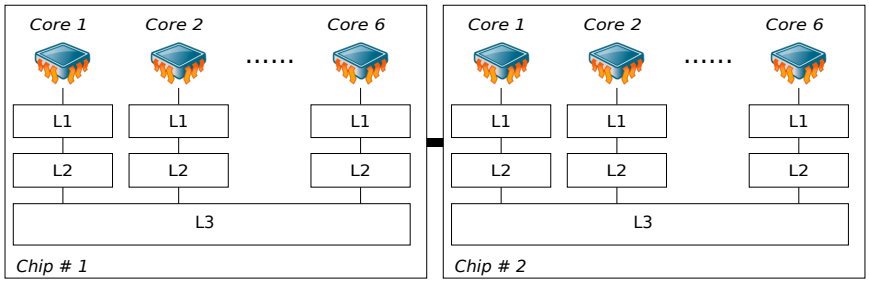


FIGURE 2.16. Architecture of a single processor (Multi Chip Module) of the Dell PowerEdge R815.

In the following we will refer the three patches we developed as:

- `original`, the original SCHED_DEADLINE implementation;
- `fmask`, SCHED_DEADLINE plus changes described in Section **??**;
- `heap`, SCHED_DEADLINE plus the heap described in Section **??**.

The reference Linux scheduler is denoted with SCHED_FIFO.

Task set generation. The algorithm for generating task sets used in the experiments works as follows. We generate a number of tasks $N = x \cdot m$, where m is the number of processors (see below), and x is set equal to 3. Similar overhead figures have been obtained with a higher number of tasks (results omitted for the sake of brevity).

The overall utilisation U of the task set is set equal to $U = R \cdot m$ where R is 0.6, 0.7 and 0.8. To generate the individual utilisation of each task, the `randfixedsum` algorithm [**ESD10**] has been used, by means of the implementation publicly made available by Paul Emberson[5]. The algorithm generates N randomly distributed numbers in $(0, 1)$, whose sum is equal to the chosen U. Then, the periods are randomly generated according to a log-uniform distribution in $[10ms, 100ms]$. The

[5]More information is available at: `http://retis.sssup.it/waters2010/tools.php`.

(worst-case) execution times are set equal to the task utilisation multiplied by the task period.

We generated 20 random task sets considering 2, 4, 8, 16, 24, 32, 40 and 48 processors. Than we ran each task set for 10 seconds using a synthetic benchmark (that lets each task execute for its WCET every period). We varied the number of active CPUs using Linux CPU hotplug feature.

Results. In Figures 2.17 and 2.18 we show the number of clock cycles required by a push operation in average, depending on the number of active cores. In Figure 2.17, we considered an average load per processor equal to $U = 0.6$, while in Figure 2.18 the load was increased to $U = 0.8$. We measured the 95% confidence interval of each average point, and it is always very low (in the order of a few tens of cycles), so we did not report it in the graphs for clarity.

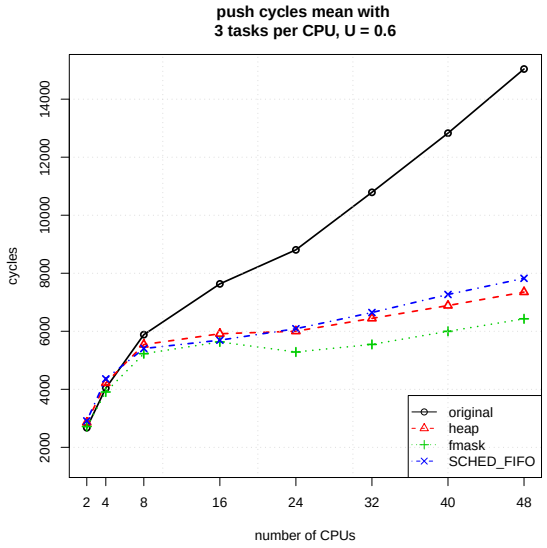


FIGURE 2.17. Number of push cycles for average loads of 0.6.

From the graphs it is clear that the overhead of the original implementation of `SCHED_DEADLINE` increases linearly with the number of processors, as expected, both for light load and for heavier load.

In `fmask`, we added the check for idle processors. Surprisingly, this simple modification substantially decreases the overhead for both types of loads, and it becomes almost constant in the number of processors. For light load, `fmask` is actually the one with lowest average number of cycles; this confirms our observation that for light loads the probability of finding an idle processor is high. For heavier loads, the probability of finding an idle processor decreases, so the `SCHED_FIFO`

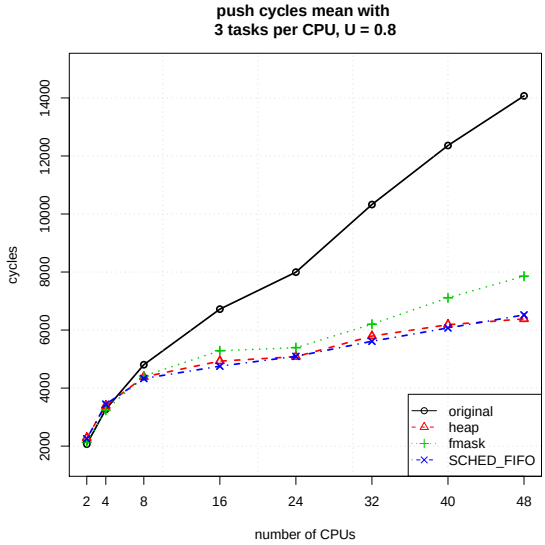


FIGURE 2.18. Number of push cycles for average loads of 0.8.

and the heap implementations are now the ones with lowest average overhead. Notice also that the latter two show very similar performance. This means that the overhead of implementing global EDF is comparable (and sometimes even lower) than implementing global Fixed Priority.

To gain a better understanding of these performance figures, it is also useful to analyse the overhead of two basic operations, enqueue and dequeue. Please remind that push and pull operations must perform at least one dequeue and one enqueue to migrate a task.

The number of cycles for enqueue operations for the four implementations is shown in Figure 2.19 for light load, and in Figure 2.20 for higher loads. The implementation with the lower enqueue overhead is `original`, because it is the one that requires the least locking and contention on shared data structures: it only requires to lock the runqueue of the CPU where the task is being moved to. `fmask` has a slightly higher overhead, as it also require to update the idle CPU mask with an atomic operation. Heap and `SCHED_FIFO` require the higher overhead as they must lock and update also global data structures (heap in the first case and priority mask in the second case). Updating the heap takes less time probably because it is a small data structure guarded by one single coarse-grain lock, whereas the prioriy mask is a complex and larger data structure with fine-grained locks. Dequeue operations have very similar performance figures and are not shown here for lack of space.

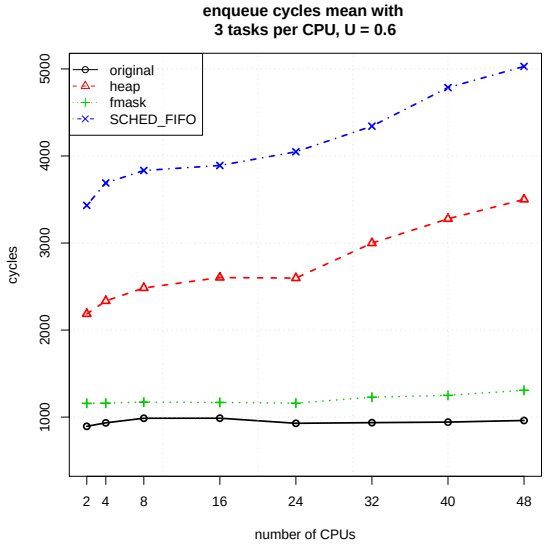


FIGURE 2.19. Number of enqueue cycles for average loads of 0.6.

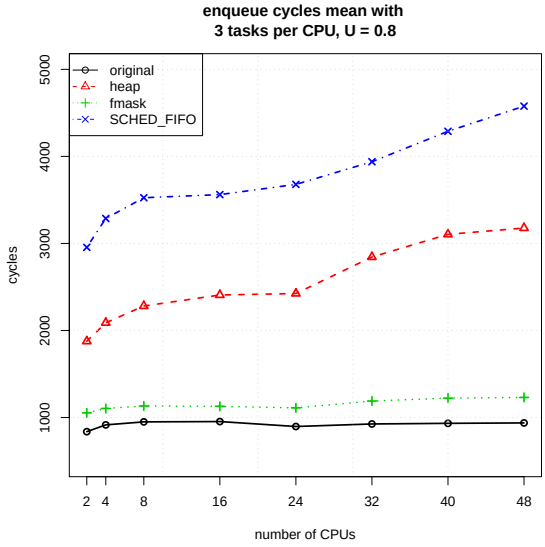


FIGURE 2.20. Number of enqueue cycles for average loads of 0.8.

Pull operations do not take advantage of any dedicated data structure. In all four schedulers, the pull operation always looks at all runqueues in sequence to find the tasks that are eligible for migration. Therefore, the execution cycles are very similar to each other. As a future work, we plan to optimize pull operations using dedicated data structures.

2.3. PRAcTISE

In this section, we propose PRAcTISE (PeRformance Analysis and TestIng of real-time multicore SchEdulers) for the Linux kernel: it is a **framework for developing, testing and debugging scheduling algorithms in user space** before implementing them in the Linux kernel, that alleviates at least part of the problems discussed above. In addition, PRAcTISE allows to compare different implementations by providing early estimations of their relative performance. In this way, the most appropriate data structures and scheduler structure can be chosen and evaluated in user-space. Compared to other similar tools, like LinSched, the proposed framework allows **true parallelism thus permitting a full test in a realistic scenario**.

The main features of PRAcTISE are:

- Rapid prototyping of scheduling data structures in user space;
- Effective, quick and extensive testing of the data structures though consistency tests;
- Real multi-core parallelism using multi-threading;
- Relative performance estimation between different algorithms and data structures in user space;
- Possibility to specify application load though probabilistic distributions of events, and statistical analysis;
- Ease of porting to the kernel or to other scheduling simulators.

PRAcTISE is available as open source software and a development version is available for download[6].

2.3.1. Developing Kernel-level code in User-Space

The wide diffusion of multi-core architectures in personal computing, servers and embedded systems, has revived the interest in multiprocessor scheduling, especially in the field of real-time applications. In fact, real-time scheduling on multi-core and multiprocessor systems is still an open research field both from the point of view of the theory and for the technical difficulties in implementing an efficient scheduling algorithm in the kernel.

[6]At the time of submission (April 29 2012, the software can be downloaded cloning the repository available at `https://github.com/Pippolo84/PRAcTISE`

Regarding the second problem, let us discuss a (non exhaustive) list of problems that the prospective developer of a new scheduler must be faced with. The task scheduler is a fundamental part of the operating system kernel: a bugged scheduler will soon crash the system, usually at random and at unexpected points. The major difficulty in testing and debugging a new scheduling algorithm derives from the fact that, when the system crashes, it is difficult to reconstruct the situation (i.e. the sequence of events and states) that led to the crash. The developer has to carefully analyse system logs and traces (for example using one of the tools described in Section 2.3.2), and reconstruct the state to understand what went wrong. More importantly, it is often impossible to impose a precise sequence of events: crashes can rarely be reproduced deterministically. Hence, it is practically impossible to run a sequence of test-cases.

This problem is exacerbated in multi-core architectures where the scheduler service routines run in parallel on the different processors, and make use of shared data structures that are accessed in parallel. In these cases, it is necessary to ensure that the data structures remain consistent under every possible interleaving of the service functions. A simple solution is to protect the shared data structure with locks. However, a single big lock reduces parallelism and performance does not scale; fine-grain locks may cause deadlock situations, without improving scalability; and lock-free algorithms are difficult to implement and prove correct. As a consequence, many important and interesting scheduling algorithms proposed in the research literature fail to be implemented on popular operating systems like Linux due to the difficulty of the task.

One reasonable approach would be to develop, debug, test and analyse the algorithms in user space. Once the main algorithm is sufficiently tested using user-space debugging and testing techniques, the same algorithm can be ported in the kernel. However, if no specific methodology is followed, the code must be written twice, increasing the possibility of introducing bugs in one of the two versions. Also, if one is unsure of which algorithm, data structure or locking strategy is more appropriate, the number of versions to implement, test, analyse by hand may become very large.

Hence, we decided to tackle the "user-space approach" by proposing a simple framework to facilitate the development, testing and performance evaluation of scheduling algorithms in user space, and minimise the effort of porting the same algorithms in kernel spaces.

2.3.2. State Of Art

Several tools exist, as open-source software, that are geared towards, or can be used as effective means to implement, debug and analyse real-time scheduling algorithms for multiprocessor systems. Each one tackles the intrinsic toughness

of this field from different angles, generally focusing on one single aspect of the problem.

A valuable tool during the development process of a scheduling algorithm would be the one that allows fast prototyping and easy debugging. Originally developed by the Real Time Systems Group at University of North Carolina at Chapel Hill, and currently maintained by P. Turner from Google, **LinSched** [**CBL+08**][7] lets developers modify the behaviour of the Linux scheduler and test these changes in user-space. One of the major strength points of this tool is that it introduces very few modifications in the kernel sources. The developer can thus write kernel code and, once satisfied by tests, it has kernel ready patches at hand. Furthermore, debugging is facilitated by the fact that LinSched runs as a single thread user-space program, that can hence be debugged with common user-space tools like GDB[8]. Even if single-threading is useful for debugging purposes, it can be a notable drawback when focusing on the analysis of behaviour assuming a high degree of concurrency. LinSched can indeed verify locking, but it cannot precisely model multi-core contention.

LITMUS$^{\mathbf{RT}}$ [**LIT**] has a completely different focus. The $LITMUS^{RT}$ patch, developed by the Real Time Systems Group at University of North Carolina at Chapel Hill, is a (soft) real-time extension of the Linux kernel that allows fast prototyping and evaluation of real-time (multiprocessor) scheduling algorithms on real hardware. The $LITMUS^{RT}$ testbed provides an experimental platform that real-time system researchers can use to simplify the development process of scheduling and synchronisation algorithms (compared to modifying a stock Linux kernel). Another nice feature of this testbed is an integrated tracing infrastructure (Feather-Trace [**BA07**]) with which performance and overhead data can be collected for off-line processing. Being a research tool rather than a production-quality system, $LITMUS^{RT}$ does not target Linux mainline inclusion nor POSIX-compliance: in other words code patches created with it cannot be seamless applied to a "vanilla" Linux kernel.

Lots of other tools exist that make kernel developers lives easier during debugging, some of them can also be used to collect performance data or even extract execution traces from a running system. Among others, these are probably part of every kernel developer arsenal:

- **KVM** [9] **+ GDB**: the very first step after having modified the kernel is usually to run it on a vitalised environment. The KVM virtual machine can here be useful as it can be attached, and controlled, by the GNU Project Debugger (GDB). However, this solution can hardly be used in

[7] v3.3-rc7 release announce: http://bit.ly/IJsyV3.

[8] http://sources.redhat.com/gdb/

[9] Kernel Based Virtual Machine: http://bit.ly/IdlzXi

presence of high concurrency; moreover, it can occasionally affect the repeatability of certain bugs.

- **perf** [**Mel10**]: the performance counter subsystem in Linux can be used to collect scheduling events and performance data from a real execution. It can also be used in conjunction with LinSched, as it can record an application behaviour that can later be played back in the simulator.
- **Ftrace** [**Ros09**]: a tracing utility built directly into the Linux kernel. Ftrace is a valuable debugging tool as it brings to Linux the ability to see what is happening inside the kernel. With the ability of synchronise a user-space testing application with kernel execution, one can track function calls up to the point where a bug may happen.
- **LTTng** [**DD06**, **Des09**]: the Linux Trace Toolkit is an highly efficient tracing tool for Linux that can help tracking down performance issues and debugging problems involving concurrent execution.

2.3.3. Architecture

In this section, we describe the basic structure of PRAcTISE. PRAcTISE emulates the behaviour of the LINUX scheduler subsystem on a multi-core architecture with M parallel cores. The tool can be executed on a machine with N cores, with N that can be less, equal to or greater than M. The tool can be executed in one of the following modes:

- testing;
- performance analysis.

Each processor in the simulated system is modelled by a software thread that performs a cycle in which:

- scheduling events are generated at random;
- the corresponding scheduling functions are invoked;
- statistics are collected.

In *testing mode*, a special "testing" thread is executed periodically that performs consistency checks on the shared data structures. In the *performance analysis mode*, instead, each thread is *pinned* on a processor, and the memory is locked to avoid spurious page faults; for this reason, to obtain realistic performances it is necessary to set $M \leq N$.

2.3.4. Ready queues

The Linux kernel scheduler uses one separate ready queue per each processor. A ready task is always en-queued in one (and only one) of these queues, even when it is not executing. This organisation is tailored for partitioned schedulers and when the frequency of task migration is very low. For example, in the case of non real-time best effort scheduling, a task usually stays on the same processor, and

periodically a load-balancing algorithm is called to distribute the load across all processors.

This organisation may or may not be the best one for global scheduling policies. For example the `SCHED_FIFO` and `SCHED_RR` policies, as dictated by the POSIX standard, requires that the m highest priority tasks are scheduled at every instant. Therefore, a task can migrate several times, even during the same periodic instance.

The current multi-queue structure is certainly not mandatory: a new and different scheduler could use a totally different data structure (for example a single global ready queue); however, the current structure is intertwined with the rest of the kernel and we believe that it would be difficult to change it without requiring major changes in the rest of the scheduler. Therefore, in the current version of PRAcTISE we maintained the structure of distributed queues as it is in the kernel. We plan to extend and generalise this structure in future versions of the tool.

Migration between queues is done using two basic functions: *push* and *pull*. The first one tries to migrate a task from the local queue of the processor that calls the function to a remote processor queue. In order to do this, it may use additional global data structures to select the most appropriate queue. For example: the current implementation of the fixed priority scheduler in Linux uses a priority map (implemented in *cpupri.c*) that records for each processor the priority of the highest priority tasks; SCHED_DEADLINE, see Section 2.2.6, uses a max heap to store the deadlines of the tasks executing on the processors.

The *pull* does the reverse operation: it searches for a task to "pull" from a remote processor queue to the local queue of the processor that calls the function. In the current implementation of SCHED_{FIFO,RR} and SCHED_DEADLINE, no special data structure is used to speed up this operation. We developed and tested in PRAcTISE a min-heap for reducing the duration of the *pull* operation, but we have not tried it yet in the kernel.

Tasks are inserted into (removed from) the ready queues using the `enqueue()` (`dequeue()`) function, respectively. In Linux, the queues are implemented as red-black trees. In PRAcTISE, instead, we have implemented them as priority heaps, using the data structure proposed by B. Brandenburg [10]. However, it is possible to implement different algorithms for queue management as part of the framework: as a future work, we plan to implement alternative data structures that use lock-free algorithms.

2.3.5. Locking and synchronisation

PRAcTISE uses a range of locking and synchronisation mechanisms that mimic the corresponding mechanisms in the Linux kernel. An exhaustive list is given in

[10]Code available here: `http://bit.ly/IozLxM`.

Table 2.3. These differences are major culprits for the slight changes needed to port code developed on the tool in the kernel 2.3.9.

It has to be noted that `wmb` and `rmb` kernel memory barriers have no corresponding operations in user-space; therefore we have to issue a full memory barrier (`__sync_synchronize`) for every occurrence of them.

Linux	**PRAcTISE**	**Action**
raw_spin_lock	pthread_spin_lock	lock a structure
raw_spin_unlock	pthread_spin_unlock	unlock a structure
atomic_inc	__sync_fetch_and_add	add a value in memory atomically
atomic_dec	__sync_fetch_and_sub	subtract a value in memory atomically
atomic_read	simple read	read a value from memory
wmb	__sync_synchronize	issue a memory barrier
rmb	__sync_synchronize	issue a read memory barrier
mb	__sync_synchronize	issue a full memory barrier

TABLE 2.3. Locking and synchronisation mechanisms (Linux vs. PRAcTISE).

2.3.6. Event generation and processing

PRAcTISE cannot execute or simulate a real application. Instead, each threads (that emulates a processor) periodically generates random scheduling events according to a certain distribution, and calls the scheduler functions. Our goals are to debug, test, compare and evaluate real-time scheduling algorithms for multi-core processors. Therefore, we identified two main events: task *activation* and *blocking*. When a task is activated, it must be inserted in one of the kernel ready queues; since such an event can cause a preemption, the scheduler is invoked, data structures are updated, etc. Something similar happens when a task self-suspends (for example because it blocks on a semaphore, or it suspends on a timer).

The pseudo-code for the task activation is function `on_activation()` described in Figure 2.21. The code mimics the sequence of events that are performed in the Linux code:

- First, the task is inserted in the local queue.
- Then, the scheduler performs a *pre-schedule*, corresponding to `pull()`, which looks at the global data structure `pull_struct` to find the task to be pulled; if it finds it, does a sequence of `dequeue()` and `enqueue()`.

- Then, the Linux scheduler performs the real schedule function; this corresponds to setting the `curr` pointer to the executing task. In PRAcTISE this step is skipped, as there is no real context switch to be performed.
- Finally, a *post-schedule* is performed, consisting of a `push()` operation, which looks at the global data structure `push_struct` to see if some task need to be migrated, and in case the response is positive, performs a `dequeue()` followed by an `enqueue()`. A similar thing happens when a task blocks (see function `on_block()`).

The pseudo code shown in Figure 2.21 is an overly simplified, schematic version of the code in the tool; the interested reader can refer to the original source code[11] for additional details.

As anticipated, every processor is simulated by a periodic thread. The thread period can be selected from the command line and represents the average frequency of events arriving at the processor. At every cycle, the thread randomly select one between the following events: `activation`, `early finish` and `idle`. In the first case, a task is generated with a random value of the deadline and function `on_activation()` is called. In the second case, the task currently executing on the processor blocks: therefore function `on_block()` is called. In the last case, nothing happens. Additionally, in all cases, the deadline of the executing task is checked against the current time: if the deadline has passed, then the current task is blocked, and function `on_block()` is called.

Currently, it is possible to specify the period of the thread cycle; the probability of an activation event; and the probability of an early finish.

2.3.7. Data structures in PRAcTISE

PRAcTISE has a modular structure, tailored to provide flexibility in developing new algorithms. The interface exposed to the user consists of hooks to functions that each global structure must provide. The most important hooks:

- `data_init`: initialises the structure, e.g., spin-lock init, dynamic memory allocation, etc.
- `data_cleanup`: performs clean up tasks at the end of a simulation.
- `data_preempt`: called each time an `enqueue()` causes a preemption (the arriving tasks has higher priority that the currently executing one); modifies the global structure to reflect the new local queue status.
- `data_finish`: *data_preempt* dual (triggered by a `dequeue()`).
- `data_find`: used by a scheduling policy to find the best CPU to (from) which push (pull) a task.
- `data_check`: implements the *checker* mechanism (described below).

[11] `https://github.com/Pippolo84/PRAcTISE`

```
pull() {
  bool found = find(pull_struct, &queue);
  if (found) {
    dequeue(&task, queue);
    enqueue(task, local_queue);
  }
}

push() {
  bool found = find(push_struct, &queue);
  if (found) {
    dequeue(&task, local_queue);
    enqueue(task, queue);
  }
}

on_activation(task) {
  enqueue(task, local_queue);
  pull();       /* pre-schedule  */
  push();       /* post-schedule */
}

on_block(task) {
  dequeue(&task, local_queue);
  pull();       /* pre-schedule  */
  push();       /* post-schedule */
}
```

FIGURE 2.21. Main scheduling functions in PRAcTISE

PRAcTISE has already been used to slightly modify and validate the global structure we have previously implemented in SCHED_DEADLINE [**?**] to speed-up `push()` operations (called *cpudl* from here on). We also implemented a corresponding structure for `pull()` operations (and used the tool to gather performance data from both). Furthermore, we back-ported in PRAcTISE the mechanism used by SCHED_FIFO to improve `push()` operations performance (called *cpupri* from here on).

We plan to exploit PRAcTISE to investigate the use of different data structures to improve the efficiency of the aforementioned operations even further. However, we leave this task as future work, since this paper is focused on describing the tool itself.

One of the major features provided by PRAcTISE is the *checking* infrastructure. Since each data structure has to obey different rules to preserve consistency among successive updates, the user has to equip the implemented algorithm with a proper checking function. When the tool is used in testing mode, the `data_check` function is called at regular intervals. Therefore, an on-line validation is performed in presence of real concurrency, thus increasing the probability of discovering bugs at an early stage of the development process. User-space debugging techniques can then be used to fix design or developing flaws.

To give the reader an example, the *checking* function for SCHED_DEADLINE *cpudl* structure ensures the max-heap property: if B is a child node of A, then $deadline(A) \geq deadline(B)$; it also check consistency between the heap and the array used to perform updates on intermediate nodes (see [**?**] for further details). We also implemented a checking function for *cpupri*: periodically, all ready queues are locked, and the content of the data structure is compared against the corresponding highest priority task in each queue, and the consistency of the flag `overloaded` in the `struct root_domain` is checked. We found that the data structure is always perfectly consistent to an external observer.

2.3.8. Statistics

To collect the measurements we use the TSC (Time Stamp Counter) of IA-32 and IA-64 Instruction Set Architectures. The TSC is a special 64-bit per-CPU register that is incremented every clock cycle. This register can be read with two different instructions: RDTSC and RDTSCP. The latter reads the TSC and other information about the CPUs that issues the instruction itself. However, there are a number of possible issues that needs to be addressed in order to have a reliable measure:

- *CPU frequency scaling and power management.* Modern CPUs can dynamically vary frequency to reduce energy consumption. Recently, CPUs manufacturer have introduced a special version of TSC inside their CPUS:

constant TSC. This kind of register is always incremented at CPU maximum frequency, regardless of CPU actual frequency. Every CPU that supports that feature has the flag *constant_tsc* in `/proc/cpuinfo` proc file of Linux. Unfortunately, even if the update rate of TSC is constant in these conditions, the CPU frequency scaling can heavily alter measurements by slowing down the code unpredictably; hence, we have conducted every experiment with all CPUs at fixed maximum frequency and no power-saving features enabled.

- *TSC synchronisation between different cores*. Since every core has its own TSC, it is possible that a misalignment between different TSCs may occur. Even if the kernel runs a synchronisation routine at start up (as we can see in the kernel log message), the synchronisation accuracy is typically in the range of several hundred clock cycles. To avoid this problem, we have set CPU affinity of every thread with a specific CPU index. In other words we have a 1:1 association between threads and CPUs, fixed for the entire simulation time. In this way we also prevent thread migration during an operation, which may introduce unexpected delays.
- *CPU instruction reordering*. To avoid instruction reordering, we use two instructions that guarantees serialisation: RDTSCP and CPUID. The latter guarantees that no instructions can be moved over or beyond it, but has a non-negligible and variable calling overhead. The former, in contrast, only guarantees that no previous instructions will be moved over. In conclusion, as suggested in [**Pao10**], we used the sequence as in Figure 2.22 to measure a given code snippet.

```
CPUID
RDTSC
code
RDTSCP
CPUID
```

FIGURE 2.22. Instruction that guarantee serialization.

- *Compiler instruction reordering*. Even the compiler can reorder instructions; so we marked the inline asm code that reads and saves the TSC current value with the keyword *volatile* .
- *Page faults*. To avoid page fault time accounting we locked every page of the process in memory with a call to `mlockall`.

PRAcTISE collects every measurement sample in a global multidimensional array, where we keep samples coming from different CPUs separated. After all simulation cycles are terminated, we print all of the samples to an output file.

By default, PRAcTISE measures the following statistics:

- duration and number of *pull* and *push* operations;
- duration and number of *enqueue* and *dequeue* operations;
- duration and number of `data_preempt`, `data_finish` and `data_find`.

Of course, it is possible to add different measures in the code of a specific algorithm by using PRAcTISE's functions. In the next section we report some experiment with the data structures currently implemented in PRAcTISE.

2.3.9. Evaluation

In this section, we present our experience in implementing new data structures and algorithms for the Linux scheduler using PRAcTISE. First, we show how difficult is to port a scheduler developed with the help of PRAcTISE into the Linux kernel; then, we report performance analysis figures and discuss the different results obtained in user space with PRAcTISE and inside the kernel.

Porting to Linux. The effort in porting an algorithm developed with PRAcTISE in Linux can be estimated by counting the number of different lines of code in the two implementations. We have two global data structures implemented both in PRAcTISE and in the Linux kernel: *cpudl* and *cpupri*.

We used the `diff` utility to compare differences between user-space and kernel code of each data structure. Results are summarised in Table 2.4. Less than 10% of changes were required to port *cpudl* to Linux, these differences mainly due to the framework interface (pointers conversions). Slightly higher changes ratio for *cpupri*, due to the quite heavy use of atomic operations (see Section 2.3.5). An example of such changes is given in Figure 2.23 (lines with a `-` correspond to user-space code, while those with a `+` to kernel code).

Structure	**Modifications**	**Ratio**
cpudl	12+ 14-	8.2%
cpupri	17+ 21-	14%

TABLE 2.4. Differences between user-space and kernel code.

The difference on the synchronisation code can be reduced by using appropriate macros. For example, we could introduce a macro that translates to `__sync_fetch_and_add` when compiled inside PRAcTISE, and to the corresponding Linux code otherwise. However, we decided for the moment to maintain the different code to highlight the differences between the two frameworks. In fact, debugging, testing and analyse the synchronisation code is the main difficulty, and the main goal of PRAcTISE;

```
[...]
-void cpupri_set(void *s, int cpu, int newpri)
+void cpupri_set(struct cpupri *cp, int cpu,
+                int newpri)
 {
-   struct cpupri *cp = (struct cpupri*) s;
    int *currpri = &cp->cpu_to_pri[cpu];
    int oldpri  = *currpri;
    int do_mb = 0;
@@ -63,57 +61,55 @@
    if (newpri == oldpri)
         return;

-   if (newpri != CPUPRI_INVALID) {
+   if (likely(newpri != CPUPRI_INVALID)) {
         struct cpupri_vec *vec =
                      &cp->pri_to_cpu[newpri];

         cpumask_set_cpu(cpu, vec->mask);
-        __sync_fetch_and_add(&vec->count, 1);
+        smp_mb__before_atomic_inc();
+        atomic_inc(&(vec)->count);
         do_mb = 1;
    }
[...]
```

FIGURE 2.23. Comparison using `diff`.

therefore, we thought that it is worth to show such differences rather than hide them.

However, the amount of work shouldered on the developer to transfer the implemented algorithm to the kernel, after testing, is quite low reducing the probability of introducing bugs during the porting. Moreover, this residual amount of handwork could be eliminated using simple translation scripts (e.g., sed). Additional macros will be introduced in future version of PRAcTISE to minimise such effort even further.

Experimental setup. The aim of the experimental evaluation is to compare performance measures obtained with PRAcTISE with what can be extracted from the execution on a real machine.

Of course, we cannot expect the measures obtained with PRAcTISE to compare directly with the measure obtained within the kernel; there are too many differences between the two execution environments to make the comparison possible: for example, the completely different synchronisation mechanisms. However, comparing the performance of two alternative algorithms within PRAcTISE can give us an idea of their relative performance within the kernel.

Results. In Linux, we rerun experiments from our previous work [**LLFC11**] on a Dell PowerEdge R815 server equipped with 64GB of RAM, and 4 AMD^R Opteron^{TM} 6168 12-core processors (running at 1.9 GHz), for a total of 48 cores. This was necessary since the *cpupri* kernel data structure has been modified in the meanwhile [12] and the PRAcTISE implementation is aligned with this last *cpupri* version. We generated 20 random task sets (using the `randfixedsum` [**ESD10**] algorithm) with periods log-uniform distributed in [10ms, 100ms], per CPU utilisation of 0.6, 0.7 and 0.8 and considering 2, 4, 8, 16, 24, 32, 40 and 48 processors. Then, we ran each task set for 10 seconds using a synthetic benchmark [13] that lets each task execute for its WCET every period. We varied the number of active CPUs using the Linux CPU hot plug feature and we collected scheduler statistics through `sched_debug`. The results for the Linux kernel are reported in Figures 2.24a and 2.24b, for modifying and querying the data structures, respectively. The figures show the number of cycles (y axis) measured for different number of processors ranging from 2 to 48 (x axis). The measures are shown in boxplot format: a box indicates all data comprised between the 25% and the 75% percentiles, whereas an horizontal lines indicates the median value; also, the vertical lines extend from the minimum to the maximum value.

In PRAcTISE we run the same experiments. As depicted in Section 2.3.6, random scheduling events generation is instead part of PRAcTISE . We varied the number of active processors from 2 to 48 as in the former case.

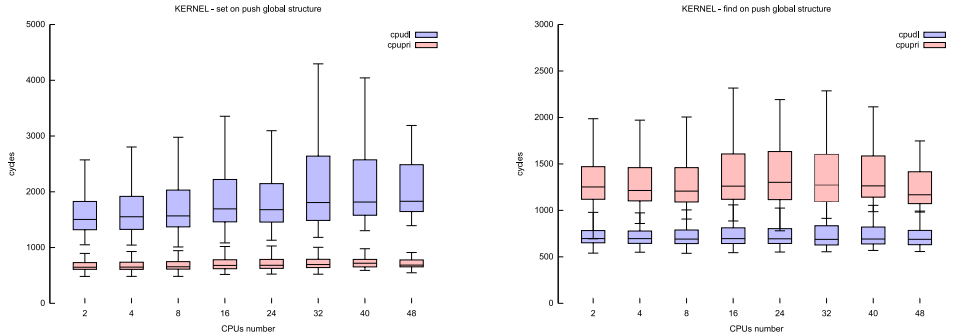
We set the following parameters: 10 milliseconds of thread cycle; 20% probability of new arrival; 10% probability of finish earlier than deadline (cpudl) or run-time (cpupri); 70% probability of doing nothing. These probability values lead to rates of about 20 task activations / (core * s), and about 20 task blocking / (core * s).

The results are shown in Figures 2.27a and 2.26a for modifying the *cpupri* and *cpudl* data structures, respectively; and in Figures 2.27b and 2.26b for querying the *cpupri* and *cpudl* data structures, respectively.

Insightful observations can be made comparing performance figures for the same operation obtained from the kernel and from simulations. Looking at Figure 2.24a

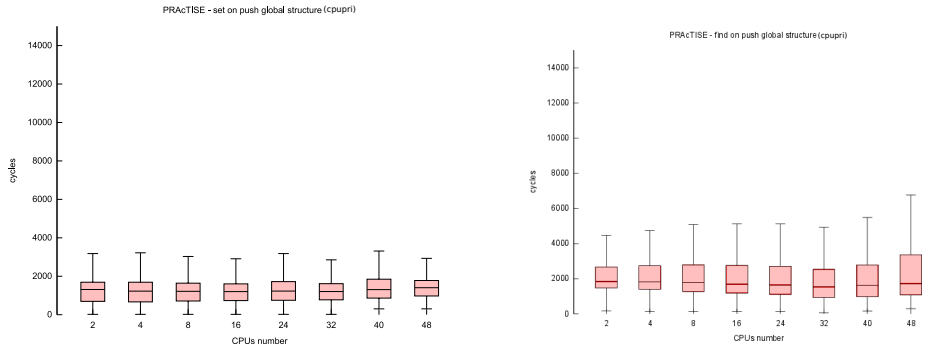
[12]More info here: `http://bit.ly/KjoePl`

[13]rt-app: `https://github.com/gbagnoli/rt-app`.



(A) modify (B) query

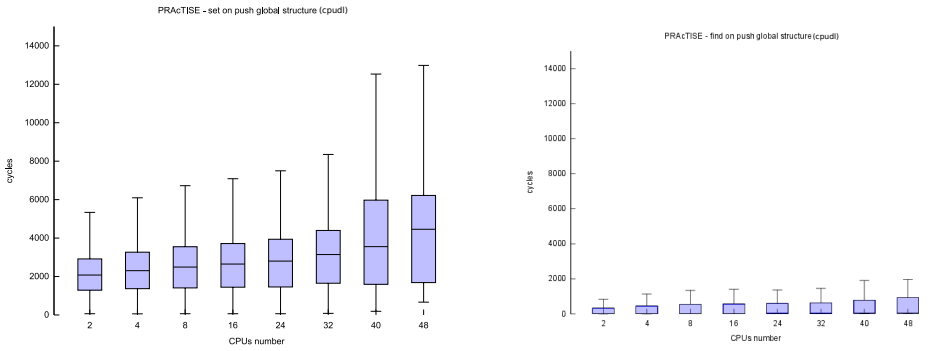
FIGURE 2.24. Number of cycles (mean) to a) modify and b) query the global data structure (*cpudl* vs. *cpupri*), kernel implementation.



(A) modify (B) query

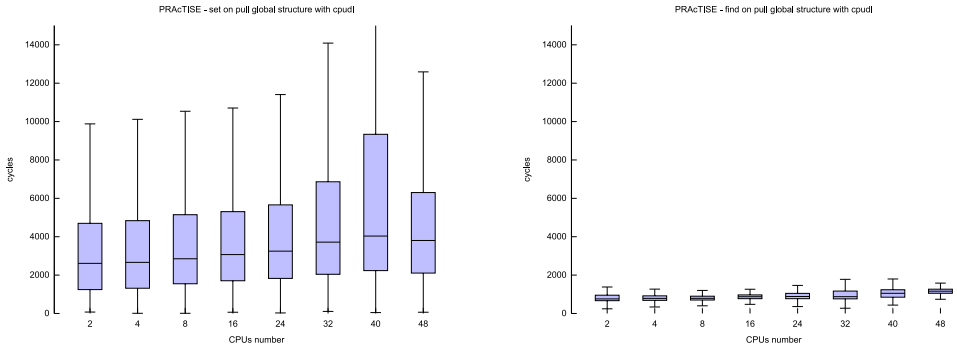
FIGURE 2.25. Number of cycles (mean) to a) modify and b) query the global data structure (*cpupri*), on PRAcTISE.

we see that modifying the *cpupri* data structure is generally faster than modifying *cpudl*: every measure corresponding to the former structure falls below 1000 cycles while the same operation on *cpudl* takes about 2000 cycles. Same trend can be noticed in Figure 2.27a and 2.26a. Points dispersion is generally a bit higher than in the previous cases; however median values for *cpupri* are strictly below 2000 cycles while *cpudl* never goes under that threshold. We can see that PRAcTISE overestimates this measures: in Figure 2.27a we see that the estimation for the *find* operation on *cpupri* are about twice the ones measured in the kernel; however, the same happens for *cpudl* (in Figure 2.26a); therefore, the relative performance of both does not change.



(A) modify (B) query

FIGURE 2.26. Number of cycles (mean) to a) modify and b) query the global data structure (*cpudl*), on PRAcTISE.



(A) modify (B) query

FIGURE 2.27. Number of cycles (mean) to a) modify and b) query the global data structure for speed-up SCHED_DEADLINE pull operations, on PRAcTISE.

Regarding query operations the ability of PRAcTISE to provide an estimation of actual trends is even more evident. Figure 2.24b shows that a *find* on *cpudl* is generally more efficient than the same operation on *cpupri*; this was expected, because the former simple reads the top element of the heap. Comparing Figure 2.27b with Figure 2.26b we can state that latter operations are the most efficient also in the simulated environment.

Moreover, we used PRAcTISE to compare the time needed to modify and query the two global data structure for push and pull operations for cpudl. As we can see in Figure 2.26a and Figure 2.26b compared against Figure 2.27a and Figure 2.27b, the results are the same, as the data structures used are the same.

We haven't compared cpudl pull operation against cpupri pull operation since the latter doesn't have a global data structure that hold the status of all run queues where we can issue find and set operations.

2.4. Conclusions

In this chapter, we first introduced a new scheduling policy for the Linux kernel that allows partitioned, clustered and global EDF scheduling with hard and soft reservations. This scheduling policy, called SCHED_DEADLINE, has been recently merged in the mainline Linux kernel.

We then reported on several enhancements we did on an old version of SCHED_DEADLINE, in order to make scheduling decisions more efficient. In particular, we detailed about an heap data structure that speeds up push decisions (where to migrate a task that has not been selected for execution by the local scheduler).

Lastly, PRAcTISE is presented. While working on further enhancing SCHED_DEADLINE performance, we felt the need to be able to fast prototype and debug quite complex mechanisms in user-space; PRAcTISE allows just to do this. Furthermore, it is possible to use the tool to perform preliminary runtime comparisons between different implementations. The tool has been used also as a validator of both SCHED_DEADLINE and SCHED_FIFO scheduling policies.

CHAPTER 3

When Theory comes to Hardware

In this chapter we build upon mechanisms described in previous sections to perform an experimental evaluation of different real-time algorithms on a real system. The experiments are run on the Linux OS and the focus is on the metrics typically of interest for developers and other people who investigate on performance issues, and not purely on schedulability analysis.

This chapter is organized as follows. Section 3.1 introduces basic concepts about memory levels and how programs execution is influenced by their presence; Section 3.2 defines this chapter contributions; Section 3.3 gives an overview of the literature on the chapter subjects; Section 3.4 describes the experimental setup on which the experimental comparison of Section 3.5 is performed; Section 3.6 concludes the chapter highlighting general considerations than can guide real-time applications developers working on modern multiprocessor machines.

3.1. Setting the Ground

Working Sets. Back in 1968 Peter J. Denning proposed the *working set model* as a mean to describe the behavior of programs in general purpose computing systems [**Den68**]. Reference system has a two-level memory structure, represented in Figure 3.1, based on the notion of *pages*. Each process in the system can only access its own private, segmented address space [1]; each segment is sliced in equal-size units, called pages. Pages can either reside in main and auxiliary memory, the two levels being connected through a bridge allowing the transfer of pages. However, a process can only work (read or write) with pages that are present in main memory; trying to access a page not currently residing there generates a page fault (the page is loaded from auxiliary memory, and usually some other page has to be swapped out to make room for the former). Roughly speaking, a working set of pages is the minimum collection of pages that has to be present in main memory for a process to work efficiently, without experiencing "unnecessary" page faults.

The *working set* of information $W(t, \tau)$ of a process at time t is formally defined to be the collection of information referenced by the process during the process time

[1]We assume here that the reader is already familiar with classical operating systems concepts, refer to Capter 7 of William Stalling's book [**Sta14**] for a thorough description.

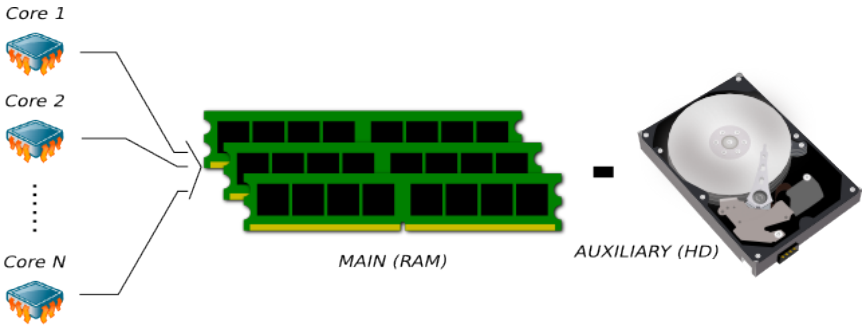


FIGURE 3.1. Two level memory system.

interval $(t-\tau, t)$. Put simply, the information a process has accessed during the last τ seconds consitutes its working set, where τ is named *working set parameter*.

As specifically stated in Denning's original paper, the model is so general that can be used for different type of information. Insted of pages, we will consider *bytes* as our unit of information. The *working set size* $\omega(t, \tau)$ is then

$$\omega(t, \tau) = number\ of\ bytes\ in\ W(t, \tau).$$

This turns to be handy, since we will deal with another level of memory, residing between main and auxiliary, called *cache memories* (see Figure 3.2).

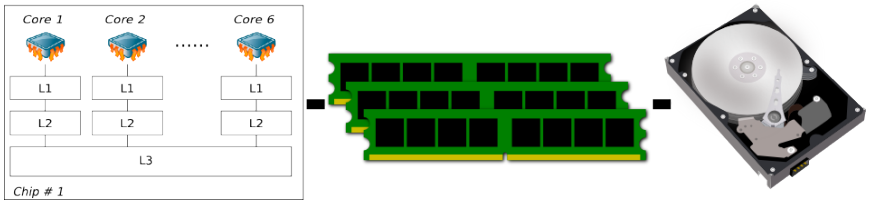


FIGURE 3.2. Three level memory system.

Cache memories. *Cache memories* are employed on modern processors systems to reduce latencies on accessing data that reside in main memory; they are fast and contains recently-accessed instructions and operands. Caches are organized in levels (or layers), see again Figure 3.2, where the fastest, and smallest, that is called *level-1* (L1), is usually complemented by deeper levels (L2, L3, etc.) being successively larger and slower. High-performance processors commonly have two separate L1 caches (*split cache*), the instruction cache and the data cache (I-cache and D-cache). On the contrary, deeper levels are usually *unified*. In multiprocessors, *shared* caches serve multiple processors (e.g., L3 in the figure), whereas *private* caches serve only one (e.g., L1 and L2 in the figure). In what follows, we use the word cache to indicate data cache for level L1 and unified cache for subsequent levels.

Caches deal with blocks of consecutive addresses called *cache lines*, where block sizes commonly range from 8 to 128 bytes. Caches are *direct mapped* if each cache line can only reside in one specific location, *fully associative* if each cache line can reside at any location or *set associative* if each line can reside at a fixed number of location. In common systems, most caches are of the last type.

Data can be read or write by processes only if residing in the first level of cache. A cache line is *useful* if is going to be referenced again. If a process references a cache line that is already present in the first level, a *cache hit* occurs. Instead, if a process references a cache line that cannot be found in a level-X cache, then it suffers a *level-X cache miss*. This can be caused by several reasons. *Compulsory misses* happen when a cache line is referenced for the first time (cache lines are loaded from main memory on-demand). Moreover, in direct mapped and set associative caches, *conflict misses* result if useful cache lines were evicted to accomodate the need of some other process (this is the major component of *preemption and migration* cost, as detailed below). *Capacity misses* are due to the fact that the WSS of a process can exceed the size of the cache (useful cache line are evicted to make room for the process own needs).

Cache effects. To better illustrate above concepts, we perform some experiments that demonstrate how the presence of cache memories can influence a process performance (what are commonly called *cache effects*). A micro-benchmark has been executed on an Intel^R Core^{TM}2 Q6600 quad-core machine. Main memory is shared among all the cores, and the cache hierarchy is on 2 levels: a private per-core 64 KB L1D (plus 64 KB L1C that we don't consider) and a global 4096 KB L2 cache. The micro-benchmark lays out an array of consecutive elements of size 64B (equal to a cache line); it then accesses them sequentially for an high number of times (to filter out eventual spikes due to external events). During this second part several quantities are measured: total access time, total L1D and L2 cache misses. These values are finally averaged on the number of time array elements were referenced.

Figure 3.3 shows average number of L1D and L2 cache misses and average time required to access an element of the array versus size of the array. The array total size is varied between 1 and 65536 KB and doubled at each step; with each element being of 64B (size of a cache line). At the beginning of each step the array is initialized and then referenced again, this time collecting measurements. Doing so, compulsory misses are ruled out. From the figure, we can see that time per step (green continuous line) remains below $5ns$ until array size becomes 4096KB, corresponing to L2 cache level size. From this point on every memory reference causes a level-L2 capacity miss (red dotted line), as L2 size for this machine is 4096KB. This is further confirmed by the steep step of L2 misses at the same point. Moreover, it is interesting to notice that L1D capacity misses (blue dashed line), that start to happen from 64KB on, seem not to influence time per step.

On modern multiprocessor machine the described behavior is common, as main memory references (last level cache misses) are much more expensive than higher levels cache misses. Considering this behavior, it is possible to argue that a process execution time shouldn't vary much, if the process only works on data that fit the highest levels of cache. We confirm this statement on the following sections.

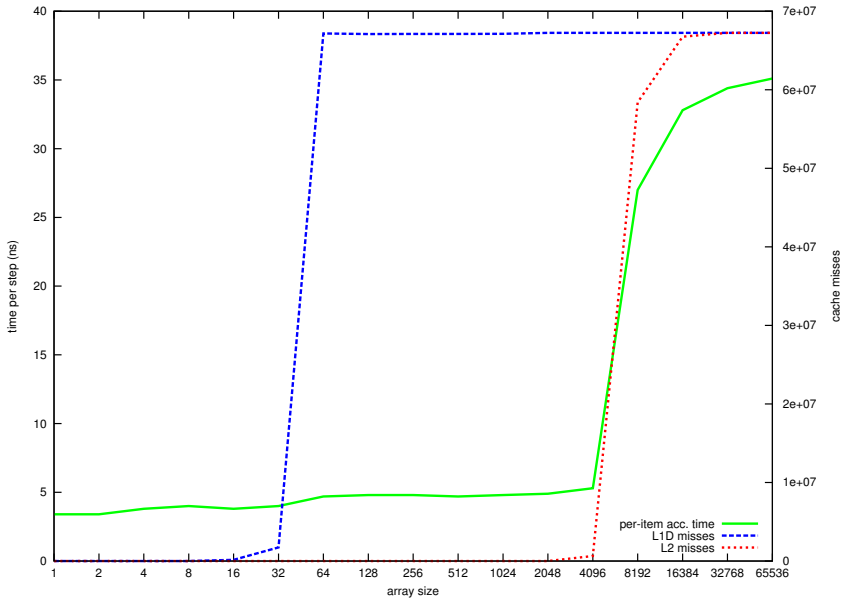


FIGURE 3.3. Cache effects on a quad-core machine.

3.2. An experimental comparison

In the following sections the performance of partitioned, clustered and global variants of Rate Monotonic (RM) and Earliest Deadline First (EDF) scheduling algorithms in the Linux OS are compared. The experimental comparison is conducted on the Linux OS due to its wide applicability (with various kernel-level patches) in the domain of real-time systems. The material has been extrapolated from the original paper by Lelli et al. [**LFCL12**].

We compare our own implementation of Global EDF (G-EDF) in the Linux kernel with respect to the fixed priority Linux scheduler (configured so as to realise RM). The goal is not to demonstrate the effectiveness of our scheduler, but rather to make a thorough performance comparison, and establish which scheduler performs better in different contexts. In order to precisely control the experiments, our methodology consists in generating sets of synthetic real-time tasks with various characteristics in terms of execution time and memory requirements and usage. The task set is then executed on a multi-core platform and the tasks' performance is

measured. The focus is on the metrics typically of interest for developers and other people who investigate on performance issues, and not purely on schedulability analysis. Indeed, we consider the laxity (tasks should not complete too close to their deadlines), the number of migrations and context switches (they may potentially affect negatively the performance), and the number and type of cache misses (they have a direct impact on the application execution times and performance). Since we focus on the comparison among different CPU scheduling policies, in this paper we only consider independent tasks. The hardware platform is a AMD^R Opteron^{TM} 6168 with 48 cores (4 sockets with 12 cores for each processor).

Our implementation of G-EDF has been made available as open-source code. This, together with the details about the configuration of the experiments, allows to reproduce and verify all the results that are presented (see the Section 3.5). Also, this allows other researchers to perform independent investigations on partitioned, clustered and global EDF scheduling on Linux, as well as to develop new schedulers and concretely compare their performance with these policies. Last, but not least, this gives to any developer the possibility to try these policies for their real-time applications.

3.3. State of Art

The comparisons available in the literature between different real-time multiprocessor scheduling solutions are almost always conducted by measuring the percentage of schedulable task sets among a number of randomly-generated ones. For example, this has been done in [**Bak06**, **BB11**, **MCF10**]. These approaches often rely on schedulability tests or simulations, and they do not involve real tasks running on a real system, thus they cannot collect such run-time metrics as the actually experienced laxity/tardiness, cache misses, context switches and migrations. For example, the possibility for these approaches to properly account for overheads due to scheduling and migration are limited. Often, these overheads are assumed to have already been considered in the Worst Case Execution Time (WCET) of the tasks. However, the scheduling policy itself may impact the WCET figures (as due to how many times a task is preempted or migrated).

Some of the main theoretical properties of EDF and RM are analysed in [**But05**], but the study refers only to uni-processor systems.

In the field of WCET analysis, in [**HP09**] a method is proposed to bound the cache-related migration delay in multi-cores, while in [**CGKS05**, **YZ08**] the focus is on devising proper task interference models. On a slightly more practical basis, memory access traces of actual program executions have been used to feed cache architectural simulators in [**MB91**,**SA04**], while in [**DCC07**,**LDS07**,**Tsa07**] some micro-benchmarks have been run on a Linux system in order to quantify the cache-related context switch delay in some specific situation (e.g., because of interrupt

processing). Although being related to our investigation, these studies concentrate on the effects that either migrations or preemptions have on task execution due to interference on shared caches in particular conditions. Instead, this study aims at a more holistic and high-level comparison of some of the available solutions for multiprocessor scheduling.

In the domain of distributed real-time scheduling on heterogeneous Grids, various schedulers have been compared by considering applications comprising of direct acyclic graphs (DAG) of computations [**TLL+11**,**SK11**] with end-to-end deadlines. Similarly, the present investigation might be expanded by comparing the performance of the schedulers under DAG-based workloads. However, due to space reasons, we do not consider DAGs, but we defer such further investigations to future research.

The line of research closer to the approach of the following sections is the one carried out by the Real-Time Systems Group at University of North Carolina at Chapel Hill. By means of their $LITMUS^{RT}$ testbed [**CLB+06**], they conducted investigations on how real overheads affect the results coming out of theoretical analysis techniques. There are several works by such group going in this direction: in [**CLB+06**] Calandrino et al. studied the behaviour of some variants of global EDF and Pfair, but did not consider fixed-priority; in [**BCA08**], Brandenburg et al. explored the scalability of a similar set of algorithms, while in [**BA09**] the impact of the implementation details on the performance of global EDF is analysed; finally, in [**BBA10b**,**BBA10a**,**A. 11**] Bastoni et al. showed intents similar to the ones of our comparison, concentrating on partitioned, clustered and global EDF on a large multi-core system. In all these works, samples of the various forms of overhead that show up during execution on real hardware are gathered and are then plugged in schedulability analysis tests, making them more accurate. However, the final conclusions about the performance of the various scheduling algorithms are actually influenced by the accuracy of the best known schedulability tests, which are often quite conservative.

Instead, in this study the tasks have just been deployed under various scheduler configurations, and their obtained performance has been systematically measured. Therefore, the conclusions drawn are simply derived from the intrinsic timing behaviour of the tasks as exhibited during the performed experiments. Further considerations regarding specifically the comparison with [**BBA10a**] follow in Section 3.6.

3.4. Experimental Setup

3.4.1. Hardware Platform

Experiments have been conducted on a Dell PowerEdge R815 server equipped with 64GB of RAM and 4 AMD^R Opteron^{TM} 6168 12-core processors (running at 1.9 GHz), for a total of 48 cores. From a NUMA viewpoint, each processor contains

two 6-core NUMA nodes and is attached to two memory controllers. The memory is globally shared among all the cores, and the cache hierarchy is on 3 levels, private per-core 64 KB L1D and 512 KB L2 caches, and a global 10240 KB L3 cache.

The hardware platform runs the Linux OS modified with the patch described in Section 2.2.

3.4.2. Task Structure

In order to compare the performance of the different scheduling strategies, it is important to precisely control the parameters of each experiment. To this end, a configurable program was developed to simulate the behaviour of a periodic real-time task. An experiment consists in running many concurrent instances of this program with different parameters.

The program parameters are the period, the expected execution time, the Working Set Size (WSS)[2], the CPU affinity[3] and the scheduler type and parameters (computation-time and period). In each periodic instance, the task accesses an array of 64 Bytes elements — i.e., equal to the size of a cache line (at all levels). Following the methodology in [**BBA10b**], the read-to-write ratio has been set equal to 4. The elements in the array are allocated contiguously and they have been accessed both sequentially and randomly, i.e. in the worst possible case with respect to data locality. In the following, we show results only for the sequential access case, however the ones for the random access case are similar in trends, but they exhibit higher numbers of cache misses. The number of elements in the array depends on the WSS of the specific round of experiments. In order to achieve actual execution times for the jobs that stay below the WCET figures (as computed by the task set generator described below), the time to access the elements of the array has been benchmarked. To impose worst-case conditions, this benchmarking phase was done with a number of tasks from 2 to 20 concurrently executing on the same CPU under `SCHED_FIFO` real-time scheduling, all of them accessing their entire WSS. In fact, each time a job is interrupted, it may resume on a different processor (due to migration); or it may resume after a long interval of time. In these cases, especially when the cumulative WSS of all the tasks exceeds the level-X cache size, it is likely that its data is not present in the cache and must be reloaded from the next level in the memory hierarchy.

3.4.3. Task Set Generation

The algorithm for generating task sets used in this paper works as follows. We generate a number of tasks $N = x \cdot m$, where m is the number of processors, and x is set equal to 2, 3 and 4. The overall utilisation U of the task set is set equal

[2]Our program activates periodically and accesses a fixed set of memory locations, which constitute the Working Set.

[3]This is the set of CPUs over which the task is allowed to run.

to $U = R \cdot m$ where R is 0.6, 0.7 and 0.8. To generate the individual utilisation of each task, the `randfixedsum` algorithm [**ESD10**] has been used, by means of the implementation publicly made available by Paul Emberson[4]. The algorithm generates N randomly distributed numbers in $(0, 1)$, whose sum is equal to the chosen U. Then, the periods are randomly generated according to a log-uniform distribution in $[10ms, 100ms]$. The (worst-case) execution times are set equal to the task utilisation multiplied by the task period. The WSS size w has been set equal to 16KB and 256KB.

The original `randfixedsum` algorithm has been used in such a way that the generated tasks resulted in a computation time higher than $C_{min} = 50\mu s$, per-task utilisation lower than U (see above) and whole task set utilisation of U itself. In fact, in the partitioned cases, we aimed at task sets able to be partitioned in a reasonably balanced way (e.g., avoiding too heavy tasks which would have forced some cores to be much more loaded than others). This was done by using the unmodified algorithm for generating tasks with a whole utilisation of m, then rescaling the resulting tasks to a whole utilisation of $U - C_{min}/T_{min}$ (with $T_{min} = 10ms$ being the minimum possible period), and in the end summing up C_{min} to all the obtained computation times.

For all the possible combinations of parameters x, R and w specified above, 3 task sets have been generated. For each task set and considered scheduling and allocation policy, an experiment of 60 seconds has been run.

3.4.4. Scheduling and Allocation

We compared two different classes of scheduling algorithms: fixed priority scheduling, with priorities assigned according to Rate Monotonic (RM, i.e. inversely proportional to the tasks' periods); Earliest Deadline First (EDF), where the priority of a task's instance is inversely proportional to its absolute deadline. Using the `SCHED_DEADLINE` scheduler, thanks to the CBS algorithm [**AB98**], the absolute deadline of each task is automatically updated by the kernel at every re-activation of the task, and set forward in time of one task period.

We also compared 3 different scheduling solutions: partitioned, clustered and global scheduling.

The reason for considering a clustered allocation is related to the underlying hardware architecture, in which different cores have non uniform access times to the system memory hierarchy. For instance, if cores that share some level of cache are placed in the same cluster, tasks running on them will likely find their working-set warmer than when migrated on some other completely unrelated core. Moreover, the interconnections between the various CPUs and the main memory banks of the

[4]More information is available at: `http://retis.sssup.it/waters2010/tools.php`.

system might entail different access latencies from a specific core to a specific set of RAM addresses.

Therefore, on the AMD^R 6168 eight clusters are considered, each containing the cores that are physically placed on the same node, i.e., that share their L3 cache. This configuration is instantiated on Linux by means of `cpusets`, a feature that makes it possible to create groups of processors and to confine tasks that belong to the set to only migrate among those CPUs (i.e., the *scheduling domain* with Linux terminology). Also, in the case of partitioned and clustered allocations, the memory allocated by tasks is *pinned* to the memory banks associated with the corresponding socket by using the same mechanism. Finally, partitioned scheduling is achieved by setting the *CPU affinity* of the tasks.

When constraining the migration capability of the tasks, i.e., in clustering and partitioning, tasks are assigned to cores so as to minimise the maximum total load on each core or group of cores. The optimum configuration has been computed off-line by solving an *integer linear programming* optimisation problem using the GNU Linear Programming Toolkit (GLPK)[5]. This results in a load which is spread as evenly as possible across the cores. However, for the partitioned scheduling cases, due to the high number of tasks, it was not possible to solve such problem optimally, but rather a maximum solving time of 960 seconds has been used, and the best solution found in that time was used.

For each allocation algorithm, both EDF and RM scheduling strategies have been considered, for a total of 6 different configurations (see Table 3.1 for easy reference).

algo/type	**part.**	**clust.**	**glob.**
EDF	P-EDF	C-EDF	G-EDF
RM	P-RM	C-RM	G-RM

TABLE 3.1. Algorithms vs. scheduling solutions: possible configurations.

3.4.5. Performance and Overhead Evaluation

The metrics against which the listed scheduling solutions are evaluated in this work are the following (further summarized in Table 3.2):

- **task migration rate**, defined as the total number of migrations occurred during a test divided by the number of tasks, provides an estimation of how frequently a task migrated;
- **task context switch rate**, defined as the total number of context switches occurred among all tasks during a test, divided by the number of tasks; provides an estimate of how frequently a task was preempted;

[5]More information is available at: `http://www.gnu.org/software/glpk/`.

- **normalised laxity**, an estimation of how big the laxity of the various tasks is, compared to the task period. For each job of each task, this is given by the actual laxity (relative deadline minus response-time) divided by the task period. The average of all the per-job values gives the normalised laxity of the task, while the average among the laxity of all the tasks in a task-set provides the normalised laxity for the task set;
- **measured utilization**, an estimation of the utilization as experimented by running each task set. For each job of each task, this is given by the actual execution time divided by the task period. The average of all the per-job values gives the utilization of the task, while the sum of the averages provides the measured utilization for the task set;
- **L1 miss rate** and **L2 miss rate**, defined as the total number of (L1 resp. L2) cache misses experienced by a job divided by the overall execution time of the job; similarly to the other metrics, per-task and per-set figures are obtained as well.

The migration and context switch rates provide insightful information about the overhead imposed by a scheduling algorithm on the system. The laxity is representative of how precisely the actual behaviour of the tasks adheres with the one that could have been expected, since it provides direct information about the exhibited response-time of the tasks. We normalise the laxity to the period in order to put it in relationship with the timing periodicity (and deadline) of the task.

Metric	Abbreviation	Definition
task migration rate	Migr. Rate	$\frac{migrations\ during\ a\ test}{number\ of\ tasks}$
task context switch rate	Cont. Switch Rate	$\frac{context\ switches\ during\ a\ test}{number\ of\ tasks}$
normalized laxity	Norm. Laxity	$\frac{relative\ deadline - resp.\ time}{task\ period}$
measured utilization	Exp. U	$\frac{experimented\ exec.\ time}{task\ period}$
L1 & L2 miss rate	L1D(L2) Miss Rate	$\frac{L1(L2)\ cache\ misses\ per\ job}{job\ exec.\ time}$

TABLE 3.2. Evaluation metrics.

In addition to this, the duration (in clock cycles) of the main scheduling and migration related functions is recorded, in order to compare the complexity of these operations for the two algorithms in the various situations. These functions are the following:

- **enqueue_task**, which inserts a task that just activated (e.g., wake-up) in the ready queue of a CPU;

- **dequeue_task**, which removes a task that just deactivated (e.g. sleep) from the ready queue of a CPU;
- **pull_tasks**, which picks ready but not running tasks from remote run-queues if they could run on the calling CPU;
- **push_tasks**, which tries to send ready tasks that are not running on the calling CPU to some other CPU where they would execute.

The performance metrics are gathered by using standard time and timer-related of the *GNU C Library*[6], by the *Hardware Performance Counters* and related tools and libraries[7] Functions durations in cycles are obtained by directly instrumenting the functions inside the kernel.

3.5. Experimental Results

Running all the tests took several days, and yielded to an extensive set of experimental data. In this section, an excerpt of such data is reported. The full obtained data set (4.3GB in compressed form) is available for download from: `http://retis.sssup.it/people/jlelli/papers/JSS2012`. Statistics come from the results of running 3 different randomly generated task sets for each configuration in terms of scheduler, allocation policy, number of tasks and their WSS.

For example, one of the 3 task sets generated in the case of U=0.6 and 96 tasks is reported in Figure 3.4. Each point corresponds to a task with period and utilisation as read from the X and Y coordinates of the point. The higher density of points for lower periods is due to the logarithmic generation of the random periods.

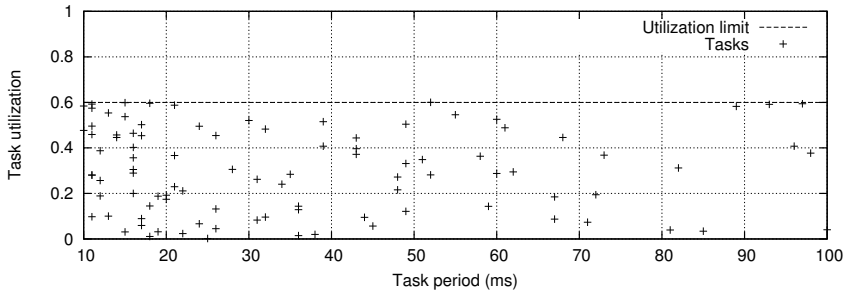


FIGURE 3.4. One of the used task-sets with U=0.6 and 96 tasks.

U	L1D	L2	L1D	L2
	WSS 16KB		WSS 256KB	
0.6	0.086	0.138	4.238	0.145
0.7	0.082	0.126	4.244	0.131
0.8	0.081	0.120	4.233	0.124

TABLE 3.3. Cache behaviour with tasks from the task-sets executed in isolation.

3.5.1. Running Each Task Alone

First, each task of each set has been executed in isolation at real-time priority (the exact real-time scheduling policy has no impact) in order to get its "reference" behaviour in terms of execution time and minimum cache misses. The obtained cache miss figures, averaged for all of the task sets, are reported in Table 3.3. The experimental variation of all the parameters is practically null due to the execution in isolation of each individual task, thus it has not been shown. As it can be seen, the cache-miss rates decrease when increasing the utilisation. In fact, longer jobs exhibit a better chance to access hot-cache data in their activations. The trends are similar for both cases of 16KB and 256KB WSS values.

3.5.2. Impact of Scheduling

Now for each configuration all the tasks in the task-set have been deployed concurrently on the platform. In the following, the focus is on the configurations with a WSS of 16KB. In all the shown results, the performance of EDF based and RM based policies are numerically and visually compared.

Table 3.4 reports the obtained normalised laxity, measured U, context switch rate, and migration rate for the various configurations: global (top table), clustered (middle table) and partitioned (bottom table) scheduling, both with EDF and RM policies. Each row corresponds to a different configuration with a total utilisation and number of tasks as indicated in the first two columns. The reported numbers constitute the average and standard deviation of the obtained metrics of interest averaged across all the tasks and the 3 task sets corresponding to each configuration.

First, we focus on how the computation times changed in the various cases, then the performance of the various scheduling algorithms from the application viewpoint.

Comparing the same configurations under EDF and RM, the obtained context switches under EDF are lower than under RM, with a tendency to have slightly lower average context switches at high overall utilisation (0.8). On the other hand,

[6]More information is available at: `http://www.gnu.org/software/libc`.

[7]More information at: `http://en.wikipedia.org/wiki/Hardware_performance_counter`, `http://icl.cs.utk.edu/papi/index.html` and `https://perf.wiki.kernel.org`.

U	#	Norm. Laxity		Exp. U		Cont. Switch Rate		Migr. Rate	
		G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM
	96	0.637 ± 0.00	0.637 ± 0.00	0.714 ± 0.00	0.714 ± 0.00	7177.260 ± 204.66	7116.566 ± 320.55	1859.906 ± 43.87	1558.247 ± 62.43
	144	0.742 ± 0.00	0.741 ± 0.01	0.755 ± 0.01	0.755 ± 0.01	6952.722 ± 351.99	6989.104 ± 205.90	2296.683 ± 182.73	1981.012 ± 225.60
0.6	192	0.790 ± 0.01	0.786 ± 0.01	0.806 ± 0.01	0.809 ± 0.01	7056.780 ± 297.99	7268.071 ± 297.61	3129.241 ± 132.60	2825.498 ± 289.64
		G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM
	96	0.586 ± 0.01	0.585 ± 0.01	0.813 ± 0.00	0.814 ± 0.00	6942.156 ± 207.28	7099.365 ± 220.57	2062.438 ± 136.11	1988.503 ± 136.62
	144	0.703 ± 0.01	0.700 ± 0.01	0.856 ± 0.00	0.859 ± 0.00	6956.086 ± 82.72	7159.338 ± 97.04	3158.907 ± 83.59	2865.542 ± 154.45
0.7	192	0.756 ± 0.01	0.560 ± 0.12	0.903 ± 0.01	0.909 ± 0.01	6689.106 ± 103.68	7112.031 ± 205.73	4365.604 ± 604.26	4201.139 ± 384.19
		G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM
	96	0.408 ± 0.08	0.057 ± 0.27	0.919 ± 0.00	0.921 ± 0.00	7212.948 ± 78.86	7549.163 ± 79.46	3329.632 ± 37.64	3471.920 ± 42.39
	144	0.246 ± 0.25	−1.277 ± 1.09	0.960 ± 0.01	0.964 ± 0.01	6447.051 ± 376.24	7209.424 ± 571.82	5261.287 ± 993.66	4944.606 ± 792.15
0.8	192	−17.268 ± 9.15	−10.330 ± 6.02	1.014 ± 0.01	1.016 ± 0.01	5879.139 ± 226.55	7146.064 ± 206.88	8088.148 ± 671.48	6886.774 ± 506.12

U	#	Norm. Laxity		Exp. U		Cont. Switch Rate		Migr. Rate	
		C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM
	96	0.658 ± 0.01	0.659 ± 0.01	0.656 ± 0.00	0.656 ± 0.00	6772.528 ± 131.22	6838.740 ± 194.84	1612.500 ± 57.60	1499.889 ± 61.64
	144	0.752 ± 0.02	0.749 ± 0.02	0.695 ± 0.01	0.695 ± 0.01	6500.192 ± 356.20	6698.641 ± 374.65	2038.775 ± 108.98	2029.683 ± 218.87
0.6	192	0.791 ± 0.02	0.787 ± 0.02	0.744 ± 0.01	0.746 ± 0.01	6550.771 ± 339.06	6979.425 ± 331.32	2600.418 ± 242.83	2571.901 ± 275.62
		C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM
	96	0.604 ± 0.02	0.603 ± 0.02	0.750 ± 0.00	0.749 ± 0.00	6478.545 ± 254.34	6684.986 ± 213.21	1844.708 ± 129.20	1860.281 ± 116.97
	144	0.706 ± 0.03	0.705 ± 0.02	0.792 ± 0.00	0.793 ± 0.00	6489.877 ± 302.18	6771.567 ± 248.40	2486.000 ± 98.06	2426.850 ± 86.98
0.7	192	0.750 ± 0.03	0.745 ± 0.03	0.838 ± 0.01	0.838 ± 0.01	6247.576 ± 86.31	6957.189 ± 127.15	2977.594 ± 160.47	3106.653 ± 148.02
		C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM
	96	0.427 ± 0.09	0.367 ± 0.11	0.850 ± 0.00	0.849 ± 0.00	6472.833 ± 156.42	7016.882 ± 119.31	2259.243 ± 40.09	2475.653 ± 14.71
	144	0.642 ± 0.05	0.220 ± 0.27	0.889 ± 0.01	0.890 ± 0.01	6045.720 ± 473.54	6751.257 ± 585.09	2995.255 ± 330.74	3135.801 ± 364.48
0.8	192	−0.017 ± 0.45	−0.470 ± 0.69	0.935 ± 0.01	0.935 ± 0.01	5834.167 ± 187.69	6868.012 ± 336.15	3766.528 ± 319.98	4011.521 ± 371.34

U	#	Norm. Laxity		Exp. U		Cont. Switch Rate		Migr. Rate	
		P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM
	96	0.599 ± 0.04	0.598 ± 0.04	0.656 ± 0.00	0.656 ± 0.00	5895.493 ± 238.37	6246.983 ± 259.01	0.000 ± 0.00	0.000 ± 0.00
	144	0.686 ± 0.05	0.684 ± 0.05	0.692 ± 0.01	0.693 ± 0.01	5575.343 ± 276.93	6006.002 ± 302.92	0.000 ± 0.00	0.000 ± 0.00
0.6	192	0.709 ± 0.06	0.707 ± 0.07	0.741 ± 0.01	0.741 ± 0.01	5560.694 ± 216.06	6132.670 ± 232.04	0.000 ± 0.00	0.000 ± 0.00
		P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM
	96	0.526 ± 0.05	0.525 ± 0.04	0.748 ± 0.00	0.748 ± 0.00	5464.538 ± 254.98	5882.622 ± 226.24	0.000 ± 0.00	0.000 ± 0.00
	144	0.617 ± 0.07	0.613 ± 0.06	0.789 ± 0.00	0.789 ± 0.00	5373.111 ± 154.76	5943.572 ± 42.84	0.000 ± 0.00	0.000 ± 0.00
0.7	192	0.633 ± 0.08	0.476 ± 0.21	0.834 ± 0.01	0.834 ± 0.00	5319.613 ± 122.86	6023.090 ± 219.59	0.000 ± 0.00	0.000 ± 0.00
		P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM
	96	0.226 ± 0.19	−0.254 ± 0.42	0.847 ± 0.00	0.849 ± 0.00	5519.233 ± 167.03	6252.608 ± 67.19	0.000 ± 0.00	0.000 ± 0.00
	144	0.537 ± 0.08	0.530 ± 0.07	0.885 ± 0.01	0.887 ± 0.01	5318.801 ± 417.72	5860.312 ± 431.40	0.000 ± 0.00	0.000 ± 0.00
0.8	192	−15.216 ± 8.93	−12.471 ± 6.76	0.929 ± 0.01	0.908 ± 0.03	4791.134 ± 184.04	5837.344 ± 170.43	0.000 ± 0.00	0.000 ± 0.00

TABLE 3.4. Statistics for the metrics of interest when WSS=16KB, under various configurations: global (top table), clustered (middle table) and partitioned (bottom table) scheduling, both with EDF and RM policies.

the migration rates of EDF based policies are significantly higher than the ones of RM based ones. However, the actual impact on the execution times of the tasks is quite limited, as it can be observed by the experimental U measured (and averaged over the 3 task sets available for each configuration) for EDF and RM based policies, which are basically equivalent.

Comparing different configurations, the context switch rate decreases with the load while, on the contrary, the migration rate increases (but decreases with the WSS, not shown). Also, note that the average experimental utilisation increases

when switching from partitioned to clustered and then to global policies, as expected due to the presence of migrations. This is also confirmed when looking at the average execution times obtained for the individual tasks. For example, Figure 3.5 shows the ratio among the execution times obtained with global, clustered and partitioned EDF scheduling, for the task set with $U = 0.6$ and 96 tasks shown also in Figure 3.4, in the two cases of 16KB and 256KB WSS. As it can be seen, in the cases of small WSS, clustered policies succeed in keeping the workload inflation at contained and negligible amounts, as compared to global strategies. However, as the WSS increases, the differences among the cache-related overheads of clustered and global scheduling tend to vanish. Working very close to the L2 cache boundary, it is very likely that each task has to reload a large part of its WSS when it is released, independently from the processor it starts running on.

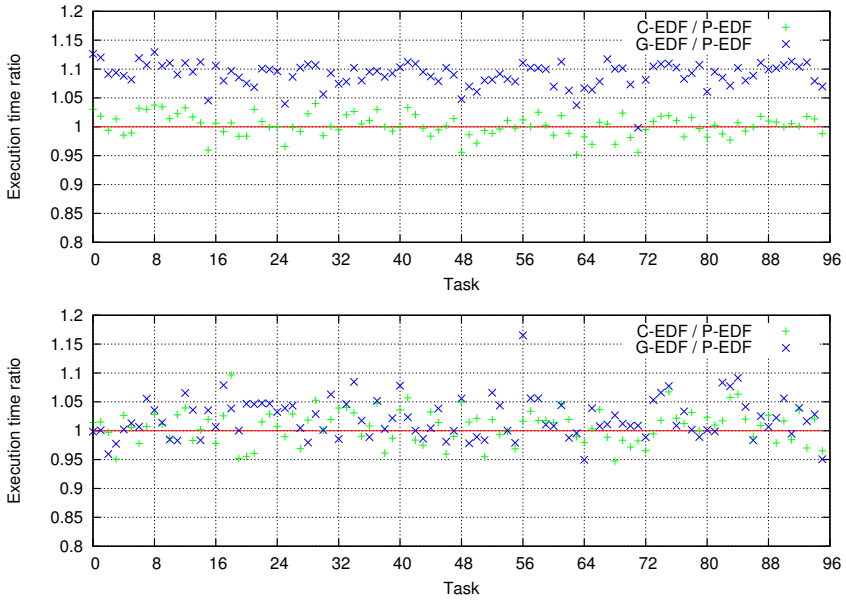


FIGURE 3.5. Execution times obtained for the various tasks (averaged over the task activations) under C-EDF (plus signs) and G-EDF (multiply signs) relative to the figures obtained under P-EDF, in the cases of 16KB (top) and 256KB (bottom) WSS.

Now, let us observe how the various scheduler configurations impact the performance as observable from within the applications themselves, i.e., let us focus on the normalised laxity figures. The average figures obtained under EDF and RM based policies are basically equivalent when the system load is low. This is due to the fact that, when a task is favoured over another by the scheduler, then the former will finish earlier and the latter will finish later, but when averaging figures this

difference is lost. The situation is different reaching the top edge with respect to utilisation (i.e., $U = 0.8$) and number of tasks. Both EDF and RM policies exhibit negative average laxity (tardiness) with 192 tasks, but only RM policies see this behaviour with also 92 and 144 tasks. It is furthermore evident that clustered policies come out to be better from this viewpoint, as slight negative values appear with 192 tasks only; moreover, C-EDF outperforms C-RM with higher average laxity figures.

In addition, let us look at the whole set of obtained laxity values, under various scheduling policies, looking at their cumulative distribution functions (CDF), first for the whole task set, then focusing on a few specific ones. Figure 3.6 shows the obtained curves when the laxity of all the tasks is considered. It is evident that (top sub-figure) global and clustered strategies (both for EDF and RM, first and second curves) differ in a negligible way from each other, and that they tend to lead to generally higher laxity values (better performance) than partitioned ones (last two curves). Also, no difference can be appreciated between EDF and RM from this viewpoint. However, zooming the figure in the critical area near a null laxity (when the tasks are close to miss their deadlines), the differences between the three policy types are more evident, with still global policies better than the others. However, partitioned EDF from this viewpoint performs much better than partitioned RM, and it achieves a similar performance to both clustered policies. This difference is due to the particular task set and the way it has been partitioned (nearly optimally – see Section 3.4.4) across CPUs. Also, it can be observed that there is a certain amount of jobs missing their deadlines. However, these are below 1% of the total jobs that have been executed, with the exception of P-RM that stands slightly above this value.

Now let us focus on the tasks with minimum period (there were 2 such tasks in the experiment with a $10ms$ period), which have the most critical timing requirements, and the ones with maximum periods (there was 1 single task with a $98ms$ period), which most often are preempted by lower period tasks. Figure 3.7 shows the obtained CDF curves when only considering the normalised laxity of these two subsets of tasks. It can be seen that global and clustered policies perform better than the partitioned ones, with RM-based policies providing an exceptionally good performance to the minimum-period tasks (Figure 3.7 top) at the expense of the maximum-period ones (Figure 3.7 bottom). These are served very badly by G-RM, while G-EDF achieves a borderline acceptable performance with nearly 18% of deadline misses. However, the reduced overheads of clustered policies (see also Section 4.4.3) show up clearly in this case, where both C-EDF and C-RM manage to perform quite well.

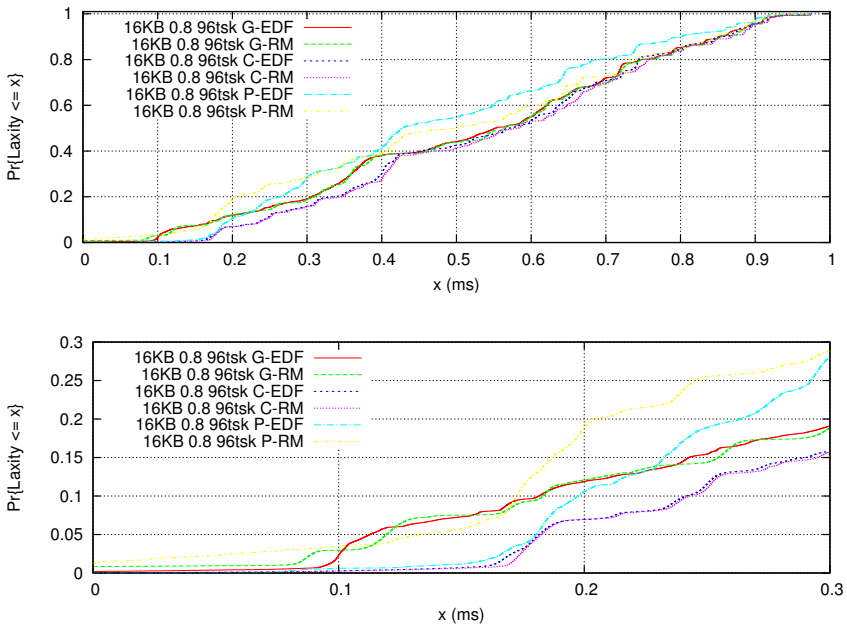


FIGURE 3.6. CDF of the normalised laxity for all the tasks as obtained under various scheduling policies, with U=0.8 and 96 tasks.

3.5.3. Working Set Size and Cache Behaviour

The impact of the WSS on the performance of the applications may be visually compared by looking at Figure 3.8. It shows the laxity CDF curves achieved on all the tasks for a WSS of 16KB (first two curves) and 256KB (last two curves), zoomed in the critical area close to the deadline miss (null laxity), obtained in the case of clustered EDF and RM scheduling, with U=0.8 and 96 tasks. Due to the higher computation times experienced by tasks with a greater WSS, the achieved laxity figures are correspondingly lower. In this plot, the difference between EDF and RM policies cannot be appreciated (see comment on Figures 3.6 and 3.7 above).

In Table 3.5 (top), the number of cache misses for L1 and L2 caches are shown, in the cases of global scheduling. They are higher than the case when tasks are run in isolation (see Table 3.3), however still not so high. Also, the difference between EDF and RM is negligible. In the middle sub-table, the figures for the cache misses of the clustered policies are reported. The numbers are slightly greater than the case of running the tasks in isolation, and very similar to the case of global scheduling (top sub-table). In the bottom sub-table partitioned policies are shown. This time the numbers are smaller than the global and clustered cases. This is due to the absence of migrations.

As expected, the cache miss rate only marginally depends on the scheduling algorithm. Rather, it has a strong dependency on the WSS, raising of one order of

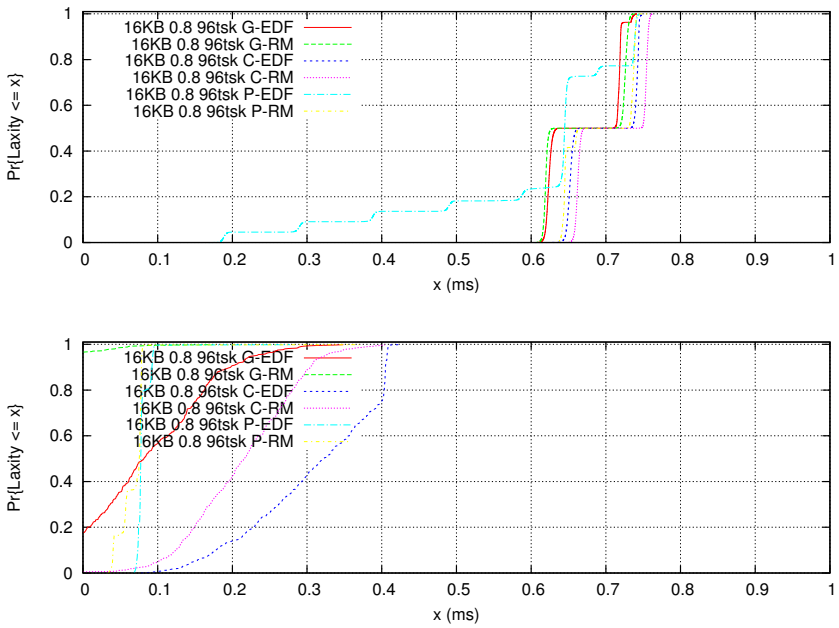


FIGURE 3.7. CDF of the normalised laxity for the minimum-period (top) and maximum-period (bottom) tasks, under various scheduling policies, with U=0.8 and 96 tasks.

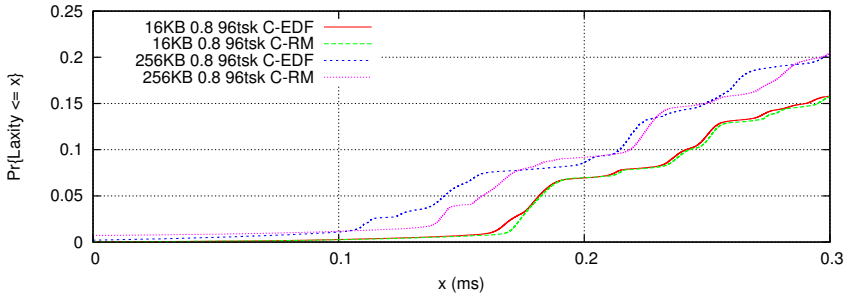


FIGURE 3.8. CDF of the obtained normalised laxity for clustered EDF and RM policies with WSS of 16KB and 256KB, with U=0.8 and 96 tasks.

magnitude when switching from a WSS of 16KB to one of 256KB. From the 16KB WSS tables, it might seem that clustering policies lead to higher cache miss rates than global scheduling, which is counter-intuitive. However, the differences are really small and sometimes within the variability range attested by the measured standard deviation. Also, for the 256KB WSS cases, partitioned policies seem to experience fewer cache misses than the others, but this is true for L1D cache miss

U	Algo.	L1D Miss Rate	L2 Miss Rate	L1D Miss Rate	L2 Miss Rate
		WSS 16KB		WSS 256KB	
0.6	G-EDF	0.100 ± 0.02	0.085 ± 0.05	4.405 ± 0.04	0.163 ± 0.07
	G-RM	0.099 ± 0.03	0.091 ± 0.05	4.392 ± 0.04	0.171 ± 0.07
0.7	G-EDF	0.101 ± 0.03	0.069 ± 0.04	4.421 ± 0.05	0.154 ± 0.06
	G-RM	0.098 ± 0.03	0.072 ± 0.04	4.352 ± 0.04	0.157 ± 0.06
0.8	G-EDF	0.119 ± 0.05	0.056 ± 0.03	4.808 ± 0.07	0.169 ± 0.07
	G-RM	0.103 ± 0.03	0.050 ± 0.02	4.344 ± 0.05	0.156 ± 0.05

U	Algo.	L1D Miss Rate	L2 Miss Rate	L1D Miss Rate	L2 Miss Rate
		WSS 16KB		WSS 256KB	
0.6	C-EDF	0.107 ± 0.03	0.101 ± 0.06	4.475 ± 0.04	0.173 ± 0.07
	C-RM	0.107 ± 0.03	0.096 ± 0.05	4.467 ± 0.04	0.178 ± 0.07
0.7	C-EDF	0.104 ± 0.02	0.083 ± 0.05	4.479 ± 0.04	0.159 ± 0.06
	C-RM	0.102 ± 0.02	0.079 ± 0.05	4.472 ± 0.04	0.165 ± 0.06
0.8	C-EDF	0.108 ± 0.03	0.064 ± 0.04	4.493 ± 0.05	0.153 ± 0.06
	C-RM	0.105 ± 0.02	0.061 ± 0.03	4.464 ± 0.04	0.163 ± 0.05

U	Algo.	L1D Miss Rate	L2 Miss Rate	L1D Miss Rate	L2 Miss Rate
		WSS 16KB		WSS 256KB	
0.6	P-EDF	0.095 ± 0.02	0.060 ± 0.03	4.504 ± 0.03	0.125 ± 0.03
	P-RM	0.093 ± 0.02	0.059 ± 0.03	4.505 ± 0.03	0.127 ± 0.03
0.7	P-EDF	0.090 ± 0.01	0.048 ± 0.02	4.529 ± 0.03	0.109 ± 0.02
	P-RM	0.088 ± 0.01	0.048 ± 0.02	4.506 ± 0.03	0.116 ± 0.03
0.8	P-EDF	0.087 ± 0.01	0.038 ± 0.02	4.787 ± 0.03	0.104 ± 0.02
	P-RM	0.088 ± 0.01	0.040 ± 0.02	4.502 ± 0.03	0.119 ± 0.03

TABLE 3.5. Cache related behaviour of global (top sub-table), clustered (middle sub-table) and partitioned (bottom sub-table) `EDF` and `RM` policies for various configurations (values are averages of all the runs for each configuration) and WSS.

rates only. Since the WSS exceeds the L1D border, little can be said about the L1D behaviour. Instead, partitioning policies lead to the lowest L2 cache miss rates as expected.

3.5.4. Scheduling Overheads Comparison

In this paragraph, the scheduling overheads obtained under various policies are compared, using the metrics described in Section 3.4.5

In Table 3.6, the overheads obtained in various configurations with a WSS of 16KB are reported. All numbers are expressed in clock cycles.

EDF based policies consistently present lower number of cycles than RM ones for enqueue operations, while, on the contrary, this is true for the partitioned case only for dequeue operations. The difference in performance is due to the completely different data structures used in the two cases. Concerning the *pull* operations, in the global and clustered scheduling cases the overheads of EDF are about twice or three times the ones of RM, whilst in the partitioned cases they are both null. Note that the small numbers shown as overheads of *pull* operations in the partitioned case are due to the few migrations at the beginning of the experiment needed to deploy the tasks over the assigned cores. Looking at the *push* operations, the overheads of

U	#	Enqueue mean time		Dequeue mean time		Push mean time		Pull mean time	
		G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM
	96	4099 ± 104.1	4867 ± 72.3	4672 ± 370.0	4518 ± 32.6	13444 ± 952.6	13895 ± 192.8	13689 ± 3772	4905 ± 124.6
	144	4147 ± 148.4	5298 ± 377.4	5158 ± 266.6	4719 ± 114.0	12637 ± 603.0	12310 ± 1018	30775 ± 4081	9143 ± 954.0
0.6	192	4078 ± 185.1	5304 ± 248.7	5803 ± 57.2	4740 ± 135.8	11938 ± 320.9	10705 ± 913.9	54308 ± 1236	15379 ± 1063
		G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM
	96	3977 ± 297.0	4810 ± 146.2	4304 ± 22.2	4183 ± 68.3	11836 ± 832.9	11743 ± 425.0	11299 ± 1398	6635 ± 859.9
	144	3840 ± 56.2	5006 ± 356.1	5288 ± 39.3	4309 ± 56.1	9674 ± 495.2	9144 ± 460.1	35216 ± 3847	13480 ± 534.9
0.7	192	3740 ± 132.8	4985 ± 220.0	5833 ± 333.4	4239 ± 83.0	8469 ± 497.5	7610 ± 478.0	64636 ± 13444	23708 ± 2175
		G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM	G-EDF	G-RM
	96	3516 ± 108.6	4295 ± 174.1	4338 ± 221.8	3796 ± 16.8	7056 ± 299.3	7250 ± 175.5	15072 ± 2540	13278 ± 131.7
	144	3367 ± 23.0	4391 ± 49.1	5317 ± 66.9	3896 ± 75.9	5428 ± 390.6	5917 ± 362.3	42396 ± 5308	24977 ± 2114
0.8	192	3420 ± 136.3	4531 ± 80.2	6172 ± 266.2	3965 ± 34.8	5830 ± 514.9	5872 ± 131.8	102436 ± 14741	40633 ± 2124

U	#	Enqueue mean time		Dequeue mean time		Push mean time		Pull mean time	
		C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM
	96	3857 ± 207.8	4961 ± 86.2	4531 ± 34.4	4552 ± 59.7	7388 ± 295.0	8840 ± 593.0	15726 ± 1143	11367 ± 1316
	144	3742 ± 98.7	5006 ± 268.1	4920 ± 75.0	4523 ± 53.6	6154 ± 195.5	7442 ± 88.4	33045 ± 2679	19974 ± 2637
0.6	192	3864 ± 164.6	5056 ± 178.3	5465 ± 147.2	4533 ± 97.6	5271 ± 408.8	6631 ± 160.8	71980 ± 12418	31313 ± 6240
		C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM
	96	3519 ± 125.7	4568 ± 192.2	4388 ± 108.2	4292 ± 23.1	6074 ± 683.1	7356 ± 152.9	17161 ± 2258	15319 ± 1646
	144	3592 ± 164.7	4487 ± 246.4	5102 ± 219.6	4185 ± 62.4	4837 ± 278.9	6712 ± 357.4	44885 ± 2780	24738 ± 1861
0.7	192	3658 ± 97.2	4925 ± 210.3	5576 ± 172.9	4377 ± 48.7	4022 ± 345.6	6144 ± 657.8	81507 ± 14115	40518 ± 2069
		C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM	C-EDF	C-RM
	96	3295 ± 34.1	4326 ± 99.5	4348 ± 127.6	3874 ± 70.8	4139 ± 217.4	6108 ± 149.2	20888 ± 1987	18963 ± 105.9
	144	3310 ± 155.6	4375 ± 309.3	5001 ± 171.6	3981 ± 121.8	3066 ± 319.7	5667 ± 236.2	50546 ± 5266	33437 ± 2427
0.8	192	3384 ± 30.2	4616 ± 105.0	5563 ± 108.3	4095 ± 81.2	2382 ± 152.2	4821 ± 138.2	90440 ± 7631	52870 ± 3211

U	#	Enqueue mean time		Dequeue mean time		Push mean time		Pull mean time	
		P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM
	96	3159 ± 225.0	4347 ± 213.3	4089 ± 57.9	5309 ± 85.7	0.0 ± 0.0	0.0 ± 0.0	131.0 ± 2.5	168.6 ± 6.2
	144	2956 ± 172.8	4106 ± 174.2	4094 ± 42.9	5397 ± 142.5	0.0 ± 0.0	0.0 ± 0.0	132.1 ± 3.2	124.2 ± 34.6
0.6	192	2751 ± 145.6	3997 ± 259.5	4014 ± 82.1	5375 ± 81.9	0.0 ± 0.0	0.0 ± 0.0	130.2 ± 7.9	121.1 ± 14.4
		P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM
	96	3471 ± 530.6	4025 ± 375.5	3990 ± 25.8	5251 ± 159.3	0.0 ± 0.0	0.0 ± 0.0	127.5 ± 4.5	151.1 ± 12.0
	144	2541 ± 171.3	3898 ± 516.0	3962 ± 157.9	5372 ± 128.4	0.0 ± 0.0	0.0 ± 0.0	126.2 ± 9.6	124.8 ± 1.2
0.7	192	2363 ± 121.5	3375 ± 152.5	3879 ± 37.9	5352 ± 106.1	0.0 ± 0.0	0.0 ± 0.0	119.0 ± 3.7	103.6 ± 10.2
		P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM	P-EDF	P-RM
	96	2754 ± 143.4	3884 ± 140.8	3847 ± 84.3	5087 ± 63.3	0.0 ± 0.0	0.0 ± 0.0	126.9 ± 5.2	96.8 ± 8.4
	144	2530 ± 29.6	3471 ± 143.4	3886 ± 44.9	5231 ± 115.4	0.0 ± 0.0	0.0 ± 0.0	115.9 ± 2.0	91.4 ± 11.1
0.8	192	2156 ± 112.6	3796 ± 536.2	3598 ± 43.5	4995 ± 88.0	0.0 ± 0.0	0.0 ± 0.0	109.1 ± 0.9	82.5 ± 2.7

TABLE 3.6. Scheduling and migration related function durations (on average, in clock cycles) for global (top sub-table), clustered (middle sub-table) and partitioned (bottom sub-table) `EDF` and `RM` policies, in the case of WSS=16KB.

G-EDF and G-RM are comparable, while C-EDF exhibits less overhead numbers than C-RM. These kind of results were expected, since in [**?**] only *push* operations were addressed, making them efficient.

3.6. Conclusions

In this chapter, an experimental comparison of various multi-processor scheduling algorithms has been performed by running synthetic workloads of real tasks on

a Linux system. The performance of the various solutions has been evaluated under diverse metrics and under multiple combinations of CPU utilisation and number of tasks.

The experimental results lead to some interesting considerations. It appears clear that global and clustered algorithms are a viable solution for multi-core platforms with a high degree of cache sharing, mainly because the overhead caused by migrations is not more costly than a "traditional" context switch. Partitioned scheduling requires a potentially costly operation of allocation of tasks to cores, and it may not manage to make use of the whole computing power available on a system. On the other hand, global scheduling has the ability to perform automatic load balancing dynamically at run-time, achieving better normalised laxity figures. However, with a large number of cores, true global scheduling may lead to a growth in overheads that needs to be kept under control, whilst clustered scheduling approaches may perform better thanks to their reduced overheads, especially when the tasks have been properly/optimally partitioned across the clusters. Indeed, at reduced sizes of the tasks WSS, clustered algorithms manage to keep applications data in the cache shared among the cores in a cluster (if present), thus achieving very little migration costs, as compared to what happens with global strategies. However, the difference tends to vanish when the WSS grows.

It is important to compare our results with the ones in [**BBA10a**], where the authors concluded that "G-EDF is never a suitable policy for hard RT systems". This may seem in conflict with our conclusions about the viability of global scheduling for RT systems. First, the reader will notice that the former conclusions concern hard real-time systems, where a formal assessment on the obtainable performance (and existence of deadline misses) can only be obtained via proper schedulability analysis techniques. On the other hand, our conclusions derive from observations of the actual behaviour of the tasks as scheduled on a real OS over a certain time horizon. In this regard, it is useful to observe that all of the task sets used in the experiments shown in our comparison have been found as non-schedulable according to one [**BC07**] of the tests for G-EDF which is known to be among the best known [**BB10**]. Second, it was not possible to set-up an experiment exactly identical to the one(s) in [**BBA10a**], due to the unavailability of the whole used data set (task WCETs, periods and clusters/partitions). Third, the analysis in [**BBA10a**] relies on the weighted schedulability, computed over task-sets with heterogeneous overall utilisation, obtained as a weighted average of the percentages of schedulable task-sets, obtained using the utilisation itself as weight. This metric was not meaningful in our context, because we were not interested in schedulability as assessed by a conservative analysis technique, but rather in the soft real-time performance as observed from the running system. On a related note, all of the task sets we considered were not schedulable by G-EDF, as stated above. Finally, another difficulty in

comparing the two approaches is due to the differences in the actual task body and in the underlying physical platform, which cause differences in the cache behaviour of the tasks when deployed on the various cores under the various configurations.

As a further remark, in both cases the conclusions cannot be considered absolutely generic, because they derive from the use of a limited number of randomly generated task sets.

In view of complete transparency, we have to state it clear that it was not possible (for practical reasons) to run the task sets till the hyper-period (even though the experiments duration was tuned so as to allow for a number of activations of each task between 600 and 6000, depending on their periods). Therefore, the reported measurements are referred to a limited observation horizon over each task-set schedule, and cannot be considered to be completely exhaustive of each and every possible foreseeable behaviour under the considered scheduling algorithms. Still, we believe the results reported in this chapter are very useful for soft real-time systems, where the main focus in on how applications may behave most of the times, and the actual worst-case that may show up once in a while may merely lead to temporary degradations of the provided service that can be easily tolerated by the system (and final users).

The implementation of G-EDF used to perform the evaluation presents acceptable overhead figures in terms of durations of the scheduling functions, being comparable with the corresponding ones present in the standard fixed priority scheduler of the Linux kernel. The performed evaluation highlights that such overheads grow anyway with the number of cores over which one is globally scheduling. *Clustered* policies lead to most of the benefits of global ones, and they can actually perform better thanks to the reduced overhead figures, posing the foundations for scalability of the technique to a high number of cores. The cost to pay is again the need for allocating tasks to clusters. However, such problem comes in a much more reduced form than the one of allocating to cores.

CHAPTER 4

Resource Reservation & Shared Resources on SMP

4.1. Introduction

In the previous chapters we detailed about our implementation of a resource reservation mechamism that can be used to provide timeliness guarantees to non-interacting real-time activities. In this chapter we start deadling with problems that arise when concurrent real-time task interacts using shared data. In this case, priority inheritance mechanims permit the donation of priorities between interacting tasks, thus allowing runtime guarantees to be provided to groups of tasks sharing data in mutual exclusion. Unfortunately, resource reservation mechanisms are incompatible with classical resource access protocols like priority inheritance. This has been shown in [**LLA04**]: if the budget is exhausted while the thread is executing inside a critical section, the priority inversion of blocked tasks can grow arbitrarily large. The authors proposed an extension of the priority inheritance protocol, called Bandwidth Inheritance BWI [**LLA04**], in which the thread that holds the lock inherits not only the priority but also the budget of the blocked tasks. The protocol has recently been extended to multi-processor systems, thus becoming the Multi-processor Bandwidth Inheritance (M-BWI) protocol [**FLC12**], and it has been implemented in LITMUS-RT [**LIT**], a research-oriented OS designed to make it easy to implement and evaluate new scheduling policies. However, LITMUS-RT is not particularly optimized for use in production systems, and it is not up-to-date with respect to current Linux OS releases.

In this Chapter we detail about an implementation of the M-BWI protocol in Linux patched with SCHED_DEADLINE. We explain the architecture and the practical aspects of our implementation. Also, we present the problems that we encountered, and a few solutions to them. Finally, we present the evaluation of the overhead introduced by the protocol.

4.2. State of the art

4.2.1. Model of a critical section

Tasks can access shared memory using mutually exclusive semaphores, often referred to as *mutex*es. The portion of code in a task between a *lock* and an *unlock*

operation on a mutex is called *critical section*. If a task needs to enter a critical section, it may be blocked by the fact that another task has already locked the corresponding mutex: this latter task is called *lock owner*, or *lock holder*. In real-time systems it is important to compute for how long, in the worst case, a task may remain blocked on a lock.

A *priority inversion* happens when a task is blocked by a low priority task. Without a proper protocol to control access to critical sections, the duration of the priority inversion may become too long, or even unbounded; for this reason, the *priority inheritance protocol* (PIP) [**SRL90**] has been proposed as a simple and effective way to reduce priority inversion. According to the PIP, when a task is blocked on a lock, the lock owner *inherits* the highest between its priority and the priorities of the blocked tasks. In this way, it cannot be preempted by medium priority tasks, thus reducing the blocking time.

Other protocols have been proposed as improvements over the PIP, notably the *priority ceiling protocol* [**GS88**], and the *stack resource policy* [**Bak91**]. However, as we will briefly discuss in the following, such protocols are not adequate for use in *open systems*.

4.2.2. Admission Control

We recall from Section 1.3.1 that one example of admission control test for the EDF+CBS algorithm on single processor systems consists in checking that the total required bandwidth does not exceed the available bandwidth: in formula, $\sum_i \frac{Q_i}{P_i} \leq 1$.

When considering also mutex semaphores and critical sections, the analysis becomes more complex, as it is now necessary to compute the maximum blocking time for each task. To compute the maximum blocking time, it is necessary to know the duration of the critical sections of all the tasks in the system. Once the blocking time has been computed, an example of admission control test for EDF is the following

$$\forall i, \sum_{D_j \leq D_i} \frac{Q_j}{P_j} + \frac{B_i}{P_i} \leq 1.$$

In an open system like Linux, however, it is not possible to ask the user to specify too many detailed information on *every task*, otherwise the system becomes too difficult to use and manage. In fact, typically in a open system tasks with different levels of criticality may coexist, and asking detailed and precise information on non-critical task may make the job of the developer/user too difficult. Also, notice that an error in the specification of the blocking times may compromise the schedulability of the whole system (i.e., any task can miss its deadline).

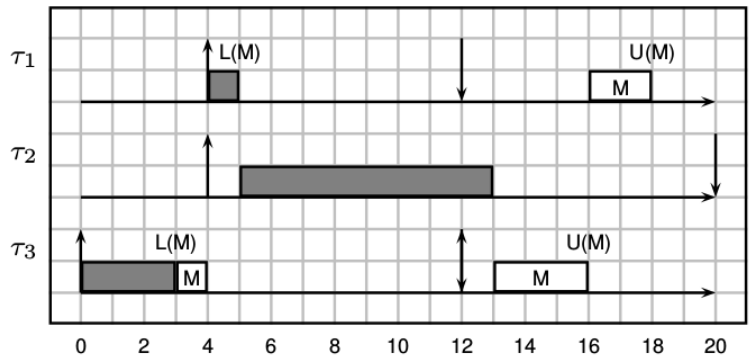


FIGURE 4.1. A task exhaust its budget while in a critical section, thus increasing the blocking time of another task.

4.2.3. Combining resource reservations and critical sections

When trying to combine resource reservations with a resource access protocol, we need to solve two problems. The first problem is concerned with how to take into account blocking time in the admission control formula, as described in the previous section.

The second problem is concerned with handling the situation that occurs when a task is in a critical section and its budget is exhausted. In that case, the scheduler algorithm *throttles* the task (i.e., suspends it) until the next period, when the budget is recharged to its maximum value. However, another task blocked on the semaphore must also wait for the budget to recharge. Therefore, the worst case blocking time may become very large.

An example of such situation is depicted in Figure 4.1, where task τ_1 suffers a long blocking time from τ_3 whose budget is exhausted at time $t = 4$ while in a critical section on mutex M.

We want to avoid this problem. At the same time, we would like to avoid the necessity to compute the blocking time of non critical tasks, and maintain the useful property of temporal isolation. So, the goals for our resource access protocol are the following

(1) We shall not require the user to specify any parameter to run the task, other that the desired budget and period (Q, P);
(2) *Temporal protection*: the performance of a task (i.e., its ability to meet its deadline) shall depend only on its parameters (Q, P), on its worst-case execution time and period, and on the duration of the critical sections of the tasks with which it interacts;
(3) If we do know the worst-case execution times (and duration of critical sections) of the task and of all *interacting tasks*, it must be possible to compute (Q, P) such that the task will meet its deadline;

(4) We want to do this on multi-core systems as well.

4.2.4. Interacting tasks

What do we mean with *interacting tasks*? In the simplest case, we say that two tasks interact when they access a critical section with the same mutex. However, since critical sections can be nested within each other, the general case is a little more complex.

A blocking chain from a task τ_i to a task τ_j is a sequence of alternating tasks and semaphores:

$$H_{i,j} = \{\tau_i \to R_{i,1} \to \tau_{i,1} \to R_{i,2} \to \ldots \to R_{i,\nu} \to \tau_j\}$$

such that τ_j is the lock owner on semaphore $R_{i,\nu}$ and τ_i is blocked on $R_{i,1}$. As an example, consider the blocking chain $H_{1,3} = \{\tau_1 \to R_1 \to \tau_2 \to R_2 \to \tau_3\}$, in which τ_3 accesses R_2, τ_2 accesses R_2 with a cs nested inside cs on R_1, τ_1 accesses R_1.

We say that a task τ_j **interferes** with task τ_i only if a blocking chain from τ_i to τ_j exists. We say that task τ_i is *independent* of, or *temporally isolated* from τ_j when there exist no blocking chain from τ_i to τ_j. We would like to maintain the temporal isolation property:

> The ability of a task to meet its deadlines depends only on its worst-case computation time and arrival pattern, its assigned budget and period, and the duration of the critical sections in the blocking chains starting from τ_i.

4.2.5. The M-BWI protocol

We informally describe here the rules of the algorithm. A more complete and detailed description can be found in [**FLC12**].

- When a task is blocked on a mutex we have several cases:
  - if the lock-owner is itself blocked, the blocking chain is followed until a non-blocked lock-owner is found;
  - if the lock-owner is executing on another processor, the blocking task actively waits for the lock owner to release the mutex;
  - if the lock-owner is not executing, then it *inherits* the budget and scheduling deadline of the blocked task.
- Therefore, when holding the lock on a mutex, a task can have a list of pairs (budget,deadline) that it can use; it will always execute consuming the budget of the earliest deadline pair.
- When a task releases the mutex, it will discard the pairs of (budget,deadline) of the blocked tasks from its list.

Rather than going through a formal analysis of the protocol, we will present here one example that demonstrate how the protocol works. In Figure 4.2 we show the schedule produced by three tasks scheduled on 2 processors with migration. All

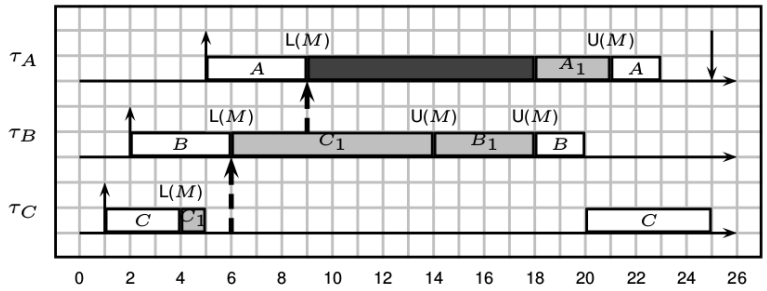


FIGURE 4.2. Example of M-BWI: τ_A, τ_B, τ_C, executed on 2 processors, that access only mutex M_1.

of them access the same mutex M_1. At time $t = 5$ tasks τ_A and τ_B are executing on the two processors, and task τ_C is the lock-owner but it is not executing. At time $t = 6$ τ_B attempts to lock the mutex, so τ_C is woken up and inherits the budget and the deadline of τ_B, while this is blocked. At time $t = 9$ also τ_A attempts to lock the mutex, and since the lock owner is already executing on another processor, it starts a spin loop actively waiting for the mutex to be unlocked. At time $t = 14$ τ_C releases the lock, and the protocol uses a FIFO policy to wake up blocked tasks, so it wakes up τ_B. Finally, when at time $t = 18$ τ_B also releases the lock, τ_A stops its active waiting and starts to execute its critical section.

Given a task τ_i, the total amount of time that other tasks execute consuming the budget Q_i and that τ_i must actively wait for a lock release, is called interference time I_i. It is possible to compute an upper bound to the interference I_i by analyzing all blocking chains starting from τ_i. In the general case of nested critical sections, the algorithm is rather complex; we remand the interested reader to [**FLC12**] for more information.

The algorithm has two important properties:

- It guarantees **temporal isolation**: a task cannot receive interference from independent tasks;
- It is possible to compute an upper bound on interference, therefore it is possible to assign the budget Q_i to a task such that it will meets all its deadlines.

4.3. Implementation

Our implementation consists of about 300 lines of codes applied on top of Linux 3.10-rc1, with SCHED_DEADLINE Version 8. Besides SCHED_DEADLINE, the work presented here is strongly based on Linux's current implementation of the Priority Inheritance (PI) algorithm. For this reason, we first proceed with a

(incomplete) review of Linux's PI design and implementation [1], and then with the detailed description of our contribution. The reader which already knows about Linux's PI infrastructure may want to skip the next subsection.

The terms *task* and *process* are used interchangeably in the rest of the document.

4.3.1. Priority Inheritance in Linux

Basic structures

In the following discussion, we will be adopting the following terminology:

Mutex.: A mutex semaphore which is shared by the processes and by which processes may interact and synchronize. The mutex structure contains a pointer to the owner of the mutex. If the mutex is not owned, this pointer is set to NULL.[2]

PI chain.: It is the same as the *blocking chain* defined in Section 4.2.4, i.e. an alternating series of mutexes and processes such that each process in the chain is blocked on the next mutex in the chain (if any), and it is the owner the previous one in the chain (if any). The PI chain causes processes to inherit priorities from a previous process that is blocked on some of its mutexes.

Lock.: A *spin-lock* that is used to protect parts of the PI algorithm. These locks disable preemption on uni-processors and prevent multiple CPUs from entering critical sections simultaneously on SMPs.

Waiter.: A structure, stored on the stack of a blocked task, that holds a pointer to the process, as well as the mutex that the task is blocked on. It also contains the node structures to place the task in the right place within the PI chains. More on this below.

Top waiter: The highest priority process waiting on a mutex.

Top PI waiter: The highest priority process waiting on one of the mutexes that a specific process owns.

Since a process may own more than one mutex, but can never be blocked on more than one, PI chains can merge at processes. Also, since a mutex may have more than one process blocked on it, but never be owned by more than one, we can have multiple chains merge on mutexes.[3]

In order to store the PI chains, Linux adopts two priority-sorted linked lists:

[1] Documentation/rt-mutex-design.txt

[2] Linux actually implements a more involved mechanism that considers "Pending Ownership" and "Lock Stealing".

[3] The maximum depth of the PI chain is not dynamic and can actually be defined by an static analysis of the code.

Waiter list.: Every mutex keeps track of all the waiters that are blocked on it by storing them in its *waiter list*. This list is protected by a lock called *wait lock*. Since no modification of the waiter list is done in an interrupt context, the *wait lock* can be taken without disabling interrupts.

PI list: Each process stores the top waiters of the mutexes that are owned by the process in its *PI list*. Note that, in general, this list does not hold all the waiters that are blocked on these mutexes. The list is protected by a lock called *PI lock*. This lock may also be taken in interrupt context, so when locking the PI lock, interrupts must be disabled.

The top of the task's PI list is always the highest priority task that is waiting on a mutex that is owned by the task. So if the task has inherited a priority, it will always be the priority of the task that is at the top of this list.

Priority adjustments

In the implementation of the priority inheritance protocol, there are several different locations in the code where a process must adjust its priority. With the help of the *PI list* it is rather easy to know what need to be adjusted; we now describe in more detail the main functions involved in this process.

The function that is responsible for adjusting the priority of a given task is `__rt_mutex_ adjust_prio`. The function first obtains the priority that the task should have, that is either the task's own normal priority, or the priority of a higher priority process that is waiting on a mutex owned by the task. By the above discussion, this is simply a matter of comparing the priority of the top PI waiter's with the task's normal priority. The function then examines the result, and if this does not match the task's current priority, the task's scheduling class methods are called to implement the actual change in priority.

Note that `__rt_mutex_adjust_prio` can either increase or decrease the priority of the task. In the case that a higher priority process has just blocked on a mutex owned by the task, the function would increase (or *boost*) the task's priority. But if a higher priority task were for some reason to leave the mutex (e.g., timeout or signal), this same function would decrease (or *unboost*) the priority of the task. This is because the PI list always contains the highest priority task that is waiting on a mutex owned by the task.

When the function `__rt_mutex_adjust_prio` is performed on a task, the nodes of the task's waiter are not updated with the new priority, therefore this task may not be in the proper locations in the waiter list of the mutex the task is blocked on and in the PI list of the corresponding owner. The function `rt_mutex_adjust_prio_chain` solves all this: it walks along the PI chain originating from the task, and updates nodes and priorities of each process it finds.

The PI chain walk can be a time-consuming operation; for this reason, `rt_mutex_adjust_prio_chain` not only defines a maximum lock depth, but it also only holds

at most two different locks at a time, as it walks the PI chain (this means that the state of the PI chain can change while in `rt_mutex_adjust_prio_chain`). When the function is called no locks is held. Then, roughly, a loop is entered where:

(1) the PI lock of the task is taken to prevent more changes to its PI list;
(2) if the task is not blocked on a mutex, the loop is exited (we are at the top of the PI chain); otherwise the wait lock of the mutex is taken to update the task's node location in the wait list;
(3) the PI lock of the task is released, and the PI lock of the mutex's owner is taken to update the task's node location in the PI list of the new process;
(4) the PI lock of the previous owner and the wait lock of the mutex are released; a new iteration of the loop is started where the previous owner will be the next task to be processed.

4.3.2. The implementation of M-BWI

The majority of our modifications (and of the difficulties we found, see below) consists in the integration of the SCHED_DEADLINE patch described in Section 2.2, with Linux's PI infrastructure described in Section 4.3.1. Indeed, at this time SCHED_DEADLINE only implements an approximated version of deadline-inheritance, in which the relative deadlines of the tasks are inherited but without control on the corresponding bandwidths; on the other hand, the implementation of PI was designed and optimized for fixed-priority tasks, hypothesis which can not be assumed for tasks scheduled with SCHED_DEADLINE policy.

In the course of our work, we have tried to keep at the minimum the modifications to original data structures in Linux and in SCHED_DEADLINE, and to maintain the original functions' semantics. For these reasons, our implementation deviates from the original M-BWI protocol as described in section 4.2.5, in that it does not consider tasks busy-waiting: like for PI, a lock-holding task inherits the scheduling parameters of a blocked task, *either* if the lock-holding task is executing or not.

Structures

As mentioned in section 4.3.1, Linux implements the waiter and the PI lists using priority-sorted linked lists. These are linked lists suitable to sort processes with fixed priorities; unlike ordinary lists, the head of this list is a different element than the nodes of the list. On the other hand, the version of deadline-inheritance which comes with SCHED_DEADLINE replaces these priority-sorted linked lists with red-black trees ordered by (absolute) deadline.

Our implementation adopts these structures to store chains of tasks and mutexes (that we will continue to call PI chains) so that it can keep the list of servers that a task can inherit sorted accordingly. By caching the left-most node of the

tree, it takes $O(1)$ to retrieve the highest priority (earliest deadline) task in the list, as for Linux's priority-sorted linked-lists.

A pointer of type `sched_dl_entity` (the scheduling entity for SCHED_ DEADLINE task) has been added to each `sched_dl_entity` structure. This pointer, named `bwi`, is required to store the effective scheduling parameters (deadline and capacity) of a SCHED_DEADLINE task, and it will point either to the task's default scheduling entity or, if a waiter with earlier deadline exists in the task's PI list, to the scheduling entity of the task's top PI waiter. In either cases, we will called the scheduling entity pointed by `bwi` the *inherited* scheduling entity or the *inherited server*.

We note that this terminology is not completely consistent with the one in [**LLA01**], even if the net result is the same. In the original presentation, each process "inherits" all the servers in its PI list, and it can execute in any of them; the CBS algorithm will then select the server (and the only runnable task within the server) with the earliest deadline. For the scope of this document, *the* inherited server, or the inherited scheduling parameters, will be this last. The difference in the terminology is mainly due to the fact that nor Linux neither SCHED_DEADLINE provide us with a server structure, and to the fact that SCHED_DEADLINE stores the server parameters in the scheduling entity of the task.

Functions

Linux resorts to the PI's logic (priority adjustments and chain walks) each time an event occurs that can possibly result in a PI chain modification. Common examples are the blocking of a task on a mutex, the release of a mutex, or the explicit call to a system call that can modify the priority and the scheduling class of a task (`sched_setscheduler`, `sched_setparam`). This remains true in the case of M-BWI for SCHED_DEADLINE tasks: priority adjustments and chain walks are necessary when the CBS algorithm modifies the deadline of a server (i.e., at capacity replenishments and and deadline updates).

For this reason, our implementation modifies the method `enqueue_task` of the SCHED _DEADLINE scheduling class, where updates and replenishments can happen, in order to fire those adjustments. Even if the logic of the chain walks and the priority adjustments remains similar to the one in Linux's PI, these modifications presented a few challenges.

IRQ safety. As mentioned in section 4.3.1, Linux does not disable interrupts before taking wait locks, because it never modifies the waiter lists in interrupt context. This is not true for SCHED_DEADLINE tasks under M-BWI, since replenishments do happen in timers interrupt context.

To solve this issue, our implementation disable interrupts before entering the corresponding critical section and re-enable them after after leaving that. Since not

all the critical section are within the scope of a functions, it was necessary to add a field in the mutex structure, in order to store the status of the interrupts.

Wake-ups in chain walks. Consider the chain $R_i \to \tau_i$ and the arrival of a new task τ_j that blocks on R_i:

$$\tau_j \to R_i \to \tau_i.$$

Linux begins updating this chain starting from τ_j. As described in section 4.3.1, τ_j's PI lock is taken to prevent additional modification to its PI list. Due to the fine-grained locking mechanism, nothing prevents τ_i from releasing the mutex a this time; suppose this is the case. Then, in the previous chain walk, R_i's wait lock is taken to insert τ_j's waiter structure in R_i's wait list, and τ_j's PI lock is released. Since R_i has no owners at this time, the chain walk needs to be interrupted and τ_j, R_i's top waiter, woken up (i.e., enqueue back to its run-queue). Finally, after the updating of τ_j scheduling parameters, the enqueue could fire a second chain walk starting from τ_j, generating a deadlock on R_i's wait lock.

Our implementation detects this situation from the status of the task to be enqueued: if the task is waking up, the enqueue will only pursue the updating on the task's deadline and capacity (if needed), without firing a (useless) chain walk.

Concurrent chain walks. The triggering of chain walks and of the corresponding priority adjustments during the enqueuing of a task, may generate other deadlock situations, in case of concurrent chain walks. This is due to the fact that the method `enqueue_task` of any Linux's scheduling class must hold the PI lock of the task that it is to be enqueued, and the run-queue's lock (the run-queue in which we are going to enqueue the task, nested inside the task's PI lock). Our implementation does not break these rules, because releasing any of these locks would have inevitably changed the semantics of the method.

To see the problem, consider a situation similar to the one above, in which a task τ_i is executing in the server inherited from τ_j (that is, τ_j has an earlier deadline than τ_i and τ_j is blocked on a mutex R_i owned by τ_i). In this situation, τ_i consumes τ_j's capacity and a replenishment is required when τ_i depletes it: in the `enqueue_task` (of task τ_i) the chain from τ_j is walked (τ_i had its scheduling parameters modified). In this case, the arrival of a new task τ_k blocking on R_i will fire a second chain walk:

$$\begin{array}{ccccc} \tau_j & \to & R_i & \to & \tau_i \\ & \nearrow & & & \\ \tau_k & & & & \end{array}$$

that can deadlock with the first:

- (in the chain from τ_j) τ_i and τ_j's PI locks held, take R_i's wait lock;
- (in the chain from τ_k) R_i's wait lock held, take τ_i's PI lock.

There are several solutions to this problem. The simplest, even if not the most rigorous one, is probably to detect the contention on one of these locks, and to just

give up after a certain number of retries. This is the solution that we adopted in our implementation, where we allow the chain walk in the enqueue to fail, eventually. Clearly, any correct solution will need to modify either the locking order (e.g., by releasing τ_i's PI lock, and so the run-queue's lock) or the locking granularity in the chain walk (e.g., by using a single lock per chain).

As already mentioned, our implementation modifies all the methods of the SCHED_DEADLINE class to use the effective scheduling parameters of the task. In particular, the method `update_curr_task`, where the execution of the current server is accounted to its capacity, now acts on the inherited server (if any). Also, the functions `push_dl_task`, `pull_dl_task`, and the methods of the `cpudl` heap structure have been modified to act on the inherited server (if any), so that global scheduling for SCHED_DEADLINE is available in SMPs ([**LLFC11**], see section 4.3.3).

4.3.3. Issues with clustered scheduling

In order to improve schedulability in multi-core systems, the Linux kernel provides a mechanism to set an affinity mask defining on which CPUs each task could execute. This approach improves performances by using application-specific information like the pattern of cache accesses or the use of specific devices. Recently, in [**GCB13**] has been demonstrated that job-level scheduling algorithms based on arbitrary processor affinity (APA) outperforms global, clustered, and partitioned approaches. However, integrate affinities inside the M-BWI protocol is not straightforward because is not yet clear what happens when a lock owner inherits the affinities of a blocked task.

Inheriting only the bandwidth and not the affinity from a blocked task could jeopardize the temporal isolation, as shown in Figure 4.3. In this case, tasks τ_1 and τ_3 can execute only on $CPU0$ while τ_2 is assigned exclusively to $CPU1$. At time $t = 2$, task τ_2 blocks on the mutex owned by τ_1, which inherits the bandwidth and the deadline continuing its execution on $CPU0$. When task τ_3 arrives at time $t = 3$ it cannot preempt τ_1 because of job priorities ($d_2 < d_3 < d_1$), but τ_2 has not been considered in the schedulability analysis of $CPU0$ thus leading to a deadline miss in $t = 11$.

Inheriting the affinity mask together with deadline and buffer is not enough to solve the problem, as shown by the example described in Figure 4.4 where tasks τ_1 can execute only on $CPU0$ while τ_2 and τ_3 are assigned exclusively to $CPU1$. At time $t = 2$, task τ_2 tries to acquire the lock owned by τ_1 which consequently migrates from $CPU0$ to $CPU1$. When the task τ_3 is activates at time $t = 3$ it preempts τ_1 which cannot execute till $t = 7$ even if its originally assigned CPU is idle, generating a deadmiss for task τ_2.

The above examples show that our implementation of M-BWI (and similarly, Linux's implementation of PI) does not extend to the case of clustered scheduling.

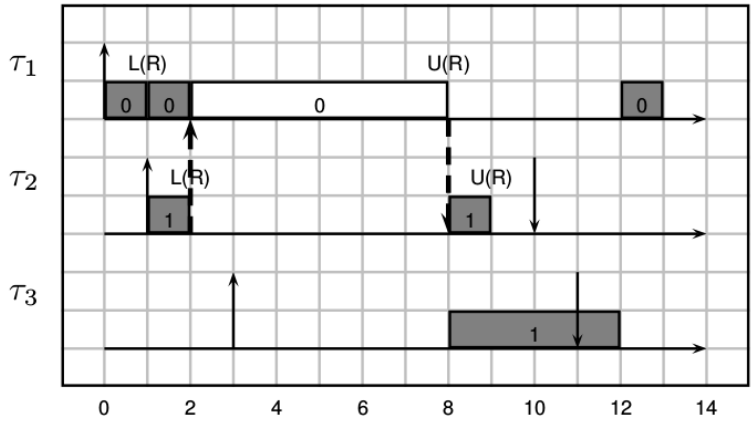


FIGURE 4.3. Deadline miss caused by a task inheriting bandwidth and deadline but not affinity from a blocked task.

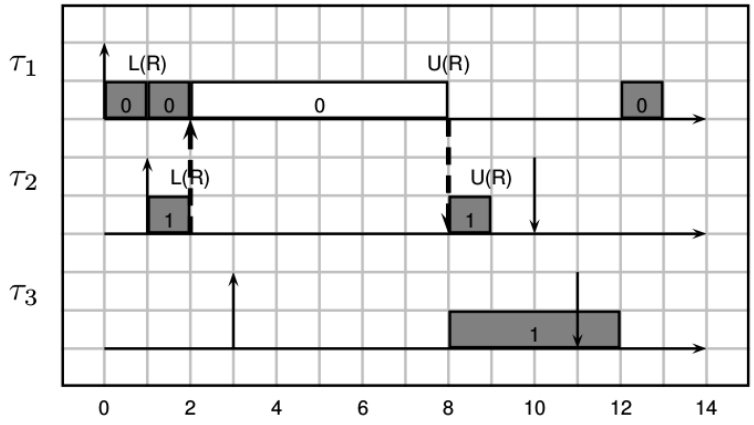


FIGURE 4.4. Deadline miss caused by a task inheriting *both* bandwidth and affinity from a blocked task.

Brandenburg [**Bra13**], proposed the Migratory Priority Inheritance protocol for clustered scheduling, and proved results of its optimality in terms of maximum PI-blocking for Job-Level-Fixed-Priority scheduling (JLFP). Under migratory priority inheritance, whenever a job J is not scheduled (but ready) and there exists a job J_i waiting for J to release a resource such that J_i is eligible to be scheduled in its assigned cluster, J migrates to J_i's cluster (if necessary) and assumes J_i's priority. The idea is that jobs should inherit (both the priority and the affinity mask) only when they "have to". However it is not yet clear how to extend this idea to the not-JLFP context, and how to implement it within the Linux kernel.

4.4. Evaluation

4.4.1. Experimental setup

We ran experiments on an Intel®Core2™ quad-core machine (Q6600) with 4GB of RAM and running at 2.4GHz.

Runtime validation consisted of executing two synthetic benchmarks. The first implements a simple situation in which two tasks share a resource protected by a mutex, and a third one, independent from the other two, is periodically activated to check if the inheritance mechanism works. The second executes a similar configuration on an SMP system.

Runtime overheads were instead measured running another benchmark, called (`rt-app`[4]). Using this application we simulated a real-time periodic load consisting of multiple SCHED _DEADLINE threads sharing resources protected by mutexes.

4.4.2. Runtime validation

We performed simple tests to validate the implementation. In the first test two threads are run that operate on the same mutex (denoted as `A`). A third thread has nothing to do with the first two, its only intent is to demonstrate that the M-BWI mechanism (when enabled) works properly. All threads are restricted to execute on the same CPU.

Figure 4.5 shows a visual representation[5] of a run when M-BWI mechanism is disabled. Threads τ_1 and τ_3 share a resource for which mutual exclusion is achieved through the use of a mutex. Both threads are periodic with periods of $24ms$ and $72ms$ respectively (deadline are set equal to periods). τ_1 executes entirely inside the critical section for $8ms$ every period, τ_3 has an execution time of $24ms$ of which $20ms$ are spent inside the critical section. τ_2 has no critical section and executes for $8ms$ every $24ms$. Thread τ_3 is the first to be activated, after a while it acquires the

[4]https://github.com/gbagnoli/rt-app

[5]Execution diagrams in this section are created through the KernelShark (https://lwn.net/Articles/425583/) utility from execution traces extracted from the kernel via `ftrace` (Documentation/trace/ftrace.txt).

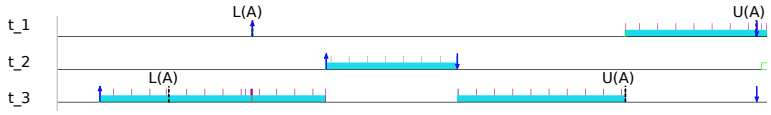


FIGURE 4.5. Two task (τ_1 *and* τ_3) sharing one resource (protected by a normal mutex). A third independent task (τ_2) arrives and preempts τ_3 even if τ_1's server has higher priority than τ_2's.

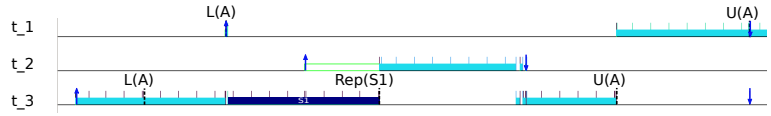


FIGURE 4.6. Two task (τ_1 *and* τ_3) sharing one resource (protected by a M-BWI-enabled mutex). A third independent task (τ_2) has to wait τ_1's server replenishment event to start executing.

mutex (`L(A)` in the figure). Then τ_1 is woken up, it tries to lock the same mutex and blocks on `A` queue, waiting for τ_3 to release it. Since it has a shorter deadline than τ_3, when τ_2 is activated it preempts τ_3 causing unexpected delay inside the critical section. τ_3 can only resume its execution once τ_2's job has finished. When τ_3 releases the mutex (`U(A)`) τ_1 executes inside the critical section and then both threads' jobs complete.

The same configuration is executed with M-BWI mechanism enabled, as depicted in Figure 4.6. In this case, when τ_1 is activated and blocks on mutex `A`, τ_3 can start executing in τ_1's server (highest priority server among τ_3 waiters), as highlighted with a darker blue shade in the figure. Since τ_3 has inherited also τ_1's deadline, when τ_2 arrives it doesn't immediately preempt τ_3. Mutex owner is actually preempted only when τ_1's server budget is exhausted and its deadline postponed (τ_2's deadline becoming the earliest), event `Rep(S1)` in the figure. After that the execution proceeds like in the previous situation.

The second test is performed on a 2 CPUs system. Two tasks (τ_1 and τ_3) share a resource protected by a M-BWI-enabled mutex `A` (we omit the standard case for brevity) and are free to execute on every CPU. Other two tasks (τ_2 and τ_4) are pinned each one on a different CPU and are independent from the others and between themselves (they are thought to create interference). Figure 4.7 zooms in a particular execution window. A job of task τ_3 arrives on CPU1 and gets scheduled, τ_3 acquires mutex `A` and enters the critical section. A few instants after a job of τ_1 arrives, τ_1 tries to acquire mutex `A` and blocks, donating its server to τ_3. The interesting part comes when τ_4 is activated. Having an earliest deadline than τ_3's original one, τ_4 should have preempted it, but its execution is delayed until τ_3

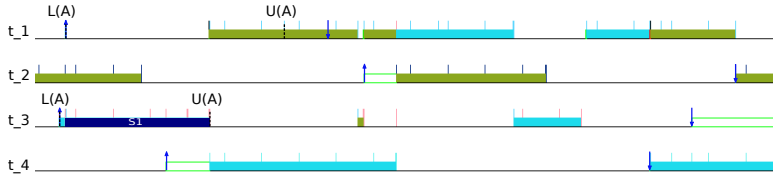


FIGURE 4.7. System with two CPUs. Two task (τ_1 *and* τ_3) sharing one resource (protected by a M-BWI-enabled mutex). Other two independent tasks (τ_2 and τ_4) are pinned each one on a different CPU.

releases mutex `A` and is consequently deboosted. After this instant of time execution continues with original parameters. Without the M-BWI mechanism working τ_3 would have been preempted inside the critical section by τ_4, delaying τ_1 execution.

4.4.3. Overheads measurements

We measured runtime overheads comparing execution of the same benchmark with the M-BWI mechanism activated, with simple deadline inheritance, and against the stock fixed priority Linux scheduler. Similarly to Brandenburg's evaluation [**Bra13**] evaluation, we launched on each core four tasks with periods $1ms$, $25ms$, $100ms$, $1000ms$ and execution times of $0.1ms$, $2ms$, $15ms$, $600ms$. The one-millisecond tasks did not access any shared resources. All other tasks shared the same lock (one lock for each core) with an associated maximum critical section length of $1ms$, and each of their jobs acquired the lock once.

We ran the task set once using the stock Linux scheduler (SCHED_FIFO with priority inheritance enabled, called `pi` in what follows), once using the original SCHED_DEADLINE implementation (deadline inheritance, `dl`) and once with the M-BWI mechanism enabled (`bwi`), for 60 seconds each. Although the same task sets can be run with priority inheritance mechanisms turned off, we don't report results coming from that configurations here as they are hardy comparable to cases when priority inheritance (or M-BWI) is enabled. In fact, execution paths inside the kernel are completely different, and unrelated functions get called, thus making the comparison of little interest for the present discussion.

Figure 4.8 shows the measurements of kernel functions, obtained using `ftrace`, that could be ill-affected by the mechanism implementation:

a) `schedule()`, scheduler core, it decides which task to run next and performs the context switch;
b) `do_futex()`, `sys_futex()` system call entry point;
c) `enqueue_task_dl()`/`enqueue_task_rt()`, enqueue a task, respectively, on the `dl` or the `rt` runqueues;
d) `rt_mutex_slowlock()`, work required to acquire a mutex;

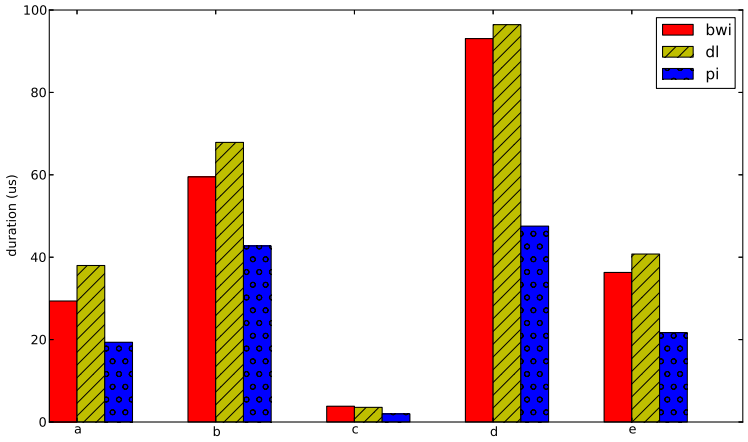


FIGURE 4.8. Kernel functions durations (in μs) from a run on a real machine.

e) `rt_mutex_slowunlock()`, work required to release a mutex.

Results show that overheads of `dl` (yellow, oblique lines, boxes) and `bwi` (red boxes) are comparable. Differences between `bwi` and `pi` (blue, small circles, boxes) measurements remain in the same order of magnitude (even if `bwi` doubles `pi` in some case). These differences can be ascribed to the slightly higher complexity of `bwi` implementation, but also to the fact that tasks interactions can be modified by scheduling the same task set using different scheduling policies (in this can have an impact on runtime overheads).

We have then modified the previous example in order to create longer PI chains: a new task with period $2000ms$ and execution time $700ms$, and two more mutexes were added to the above task set. Like in the previous example, there is a task that does not use any resource; no task accesses more than two mutexes, but the resulting PI chain can reach a depth of 4:

$$\tau_1 \rightarrow R_1 \rightarrow \tau_2 \rightarrow R_2 \rightarrow \tau_3 \rightarrow R_3 \rightarrow \tau_4$$

We replicated this task set 3 times for a total of 15 tasks, due to bandwidth constraint. The results displayed in Figure 4.9 show that the effect of the chain's depth contributes in an equivalent amount for the three implementations.

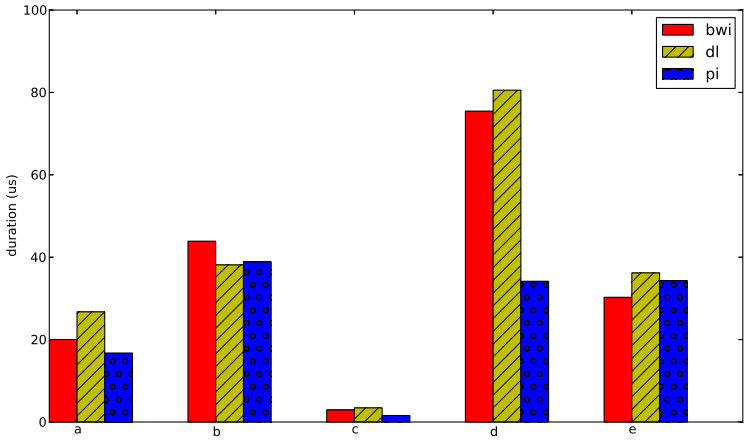


FIGURE 4.9. Kernel functions durations (in μs) with nested critical sections, from a run on a real machine.

4.5. Conclusions

In this Chapter we presented an implementation of the M-BWI protocol in the Linux kernel with the SCHED_DEADLINE patch. We tried to be as adherent as possible to the original implementation of the priority inheritance protocol in Linux and to the SCHED_DEADLINE patch by minimizing the number of modifications. The overhead of our implementation are comparable with the typical overhead of SCHED_DEADLINE and slightly larger than the overhead of the PI with SCHED_FIFO.

As future work, we are investigating the problems that we encountered when trying to inherit the affinity mask of the lock-owner task. We believe that to overcome such difficulties it is necessary to rethink the current implementation of the SCHED_DEADLINE patch, by introducing the concept of *server* as a separate scheduling entity in the implementation.

CHAPTER 5

Conclusions

In this last chapter we briefly review the results achieved while working at this thesis, suggesting possible directions for future work.

5.1. Summary of Results

The proposed goal of this thesis was to reduce the gap between real-time literature and industry, applying and implementing real-time scheduling mechanisms on General Purpose Operating Systems. The focus has been on the Linux Operating System in the context of modern Symmetric Multiprocessing computing platforms.

We started providing a general introduction, notation and definitions concerning real-time systems in Chapter 1. We also detailed about differences and peculiarities of Uniprocessor and Multiprocessors real-time sytems. We then concluded the first Chapter with a taxonomy of different approaches in designign Operating Systems and the predictability issues that may arise when such systems are not originally thought for real-time system (as Linux is).

In Chapter 2, we described the structure of the Linux scheduler, and detailed about our implementation of the global Earliest Deadline First scheduling algorithm with hard and soft resource reservations (Constant Bandwidth Server). Such implementation is now part of the Linux scheduler as a scheduling policy, and goes under the name of SCHED_DEADLINE. We also described several improvements we achieved, over the original version of the scheduling policy, in performing efficient global scheduling decisions. Chapter 2 contains also a detailed description PRAcTISE (Performance Analysis and Testing of real-time multicore Schedulers). PRAcTISE is a framework we created that can be used to ease developing, testing and debugging scheduling algorithms in user space, before implementing them in the kernel.

We then used the work of Chapter 2 to perform, in Chapter 3, an experimental comparison of the performance of partitioned, clustered and global variants of Rate Monotonic (RM) and Earliest Deadline First (EDF) scheduling algorithms in the Linux Operating System. The purpose was to highlight where each of the two algorithms excels or what are instead the problems that real-time application developers may encounter in using such scheduling algorithms.

We then open the analysis to set of tasks that share resources. Given the fact that resource reservation mechanisms are incompatible with classical resource access protocols, we detailed about our implementation of the Multiprocessor Bandwidth Inheritance protocol within SCHED_DEADLINE. The protocol levers on the basic idea of classical priority inheritance and extends it to work when interacting tasks are scheduled using the Constant Bandwidth Server mechanism. It basically enables donation of computing capacity between tasks accessing resources protected by mutexes. We detail about technicalities of the implementation and we perform an experimental evaluation of performance.

5.2. Future Work

There are several directions for future work and for refining and extending the results presented in this thesis, as we detail next.

Hierarchical Scheduling

Hierarchical scheduling is an interesting methodology for design and deploying real-time application, since it enables component-based design and analysis, and supports temporal isolation among competing components. In this thesis we only detailed about configurations where each task is assigned a single reservation (we can also say that there is a 1-to-1 relashionship between tasks and reservations). While this approach gives an high level of granularity in specifying real-time application requirements, it can be not easy to translate requirements of an application composed by several concurrent activities in single reservations for each of the composing elements (or actually impossible, e.g., legacy applications of which we could not know the internal design). It is therefore interesting to extend the presented mechasims to provide hierarchical scheduling.

Resource Sharing

The meachanisms described in Chapter 4 are a basic extension of classical priority inheritance to systems that implement resource reservation mechanisms. As we already stated, what we implemented is a first attemp in having the Multiprocessor Bandwidth Inheritance protocol working on top of our implementation of Global EDF scheduling. While providing good runtime performace, our first attemp is however quite far from a proper solution to the problem, both from a theoretical and practical point of view (for example, as discussed in Section 4.3.2, our implementation doesn't consider busy-waiting of task that blocks on a mutex). We are currently studying how to extend that implementation to actually remove the need of busy-waiting, thus simplyfing practical implementation.

Reclaiming Mechanisms

The common practise in using resource reservation mechanisms, while guaranteeing applications performance, is to over-reserve resources based on worst-case

resource usage estimates. This may lead to unnecessary wasting of system resources, given that, especially for soft real-time systems, application requirements are usually way below worst-case scenarios, and nothing terrible happens if some deadlines are missed from time to time. Resource Reclaiming mechanisms allow sharing of the over-allocation of resources. We are currently working on modifying the SCHED_DEADLINE scheduling policy to implement the GRUB (Greedy Reclamation of Unused Bandwidth) scheduling algorithm. As the required modifications are quite simple and minimal, the benefits given by such an effort are likely to be really promising.

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