145071 - Real time operating systems and middleware

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Contents

Introduction to the Course

Material

- Slides available from moodle
- Interested students can have a look at: *Giorgio Buttazzo*, **HARD REAL-TIME COMPUTING SYSTEMS: Predictable Scheduling Algorithms and Applications**

Exam

Written Exam

- 3 questions, 30 minutes per question
- Each answer gets a score from 0 to 30
- (Optional) project

Oral Exam

- Discussion of the written exam
- Open Questions or Discussion on a project

Prerequisites

- Programming skills: C, maybe C++.
  You must know how ot code in C (optionally C++). This is not about knowing the C syntax, it is about writing good and clean C code.
  To help overcome this lack of prerequisites please consider reading the book *Kerrigan & Ritchie*,**The C Programming Language**
- Knowledge about Operating Systems.
  This prerequisites is met if you have taken the course *Sistemi Operativi 1* or similar exams.
  Alternatively please refer to a good Operating Systems book (e.g. Stallings,...).
  This includes how to use a shell, basic POSIX commands, `make`, how to compile, ...

Overview of the Course

The course will cover 6 main macro areas of real time operating systems and middleware.

1. Real Time Systems:

   - Real-Time Computing and temporal constraints.
     Real time systems are software and hardware systems (hence computing systems), that have to comply with temporal constraints.
   - Definitions and task model
     We will make things much clearer and better defined by introducing a sequence of definitions and mathematical models that will allow us to given this notion of temporal constraint a well founded meaning.

- Real-Time Scheduling
  We will also study solutions that allow us to enforce these real time constraints and this solution will have much to do on how we schedule shared resources.

2. Real-Time programming, RT-POSIX, pthreads,...
   We will move to a concrete ground and see what is the exact shape that these notions take once they are moved in a computer program.
3. Real-Time Scheduling algorithms:

   - Fixed Priority scheduling, RM, DM
   - EDF and dynamic priorities
   - Resource Sharing (Priority Inversion, ...)

   As regards the Real-Time scheduling we will see many interesting policies, but since this is not a course on Real Time scheduling what we will do is provide the knowledge of real time scheduling so that the reader will be able to understand the mechanism of real time operating systems and thereby make best use of these technologies in future projects.
4. Operating System Structure

   - Notes about traditional kernel structures
     In order to keep latencies in check, we need proper technological solutions that make our operating systems differ quite a bit from standard operating systems.
   - Sources of kernel latencies
   - Some approaches to real-time kernels (e.g. dual kernel approach, interrupt pipes, micro-kernels, monolithis kernels and RT)

5. Real-Time Kernels and OSs.
6. Developing Real-Time applications

Real-Time Operating Systems

In order to discuss about the Real-Time systems we need to provide some basic definitions:

Definition 1: Real-Time Operating Systems (RTOS)

Operating Systems that provide support to Real-Time Applications

Definition 2: Real-Time application

the correctness depends not only on the output values, but also on the time when such values are produced

Definition 3: Operating Systems (OS)

- Set of computer programs, of critical programs to be precise: because they have to be written efficiently, otherwise the hardware resources get disrupted, hence the system cannot operate correctly.
- Interface between applications and hardware.
  Whenever an application interacts with an hardware, it is not of the developer interest to directly control the hardware. The Operating System provides an API that enables you to open a connection to a peripheral and takes care of all the low level interactions. On this regard, understanding the notion of interrupt will be of fundamental importance, because it is, essentially, what gave rise to concurrent programming: in the case we would like to interact with a peripheral, rather than continuously check if the peripheral has ended what it is supposed to do, you can tell the peripheral to communicate when

it has completed the given task.
Anyway the Operating systems acts as an interface towards the hardware and hides away all these complex details.
- Control the execution of application programs
- Manage the hardware and software resources

Since the Operating System is something that lies in-between the user application and the hardware resources we can summarize the aforementioned interpretation of

- **Service Provider** for user programs (i.e. exports a programming interface). Service Provider
This concept looks at the OS from the perspective of the software application, in the sense that the Operating Systems provides to the application a series of services:

  – Process Synchronization mechanism
  – Inter-Process Communication (IPC)
  – Process/Thread Scheduling, i.e. ways to create and schedule tasks
  – Input/Output
  – Virtual Memory

  And all these services are accessible through an API.
- **Resource Manager** Resource Manager
If you think at the Operating System as a Resource Manager, then it is something that takes care of many things:

  1. **Process Management** Process Management
  The fact that multiple applications can run at the same time on a PC, even though there is a small amount of processor available to manage these applications. (generally 2,4 or 8).

  The number of application that you are likely to create is often on the hundreds, hence it is necessary to make an appropriate sharing of the limited resources that you have in order for all the applications to live correctly.

  2. **Memory Management** Memory Management
  Supposing one is using a 64-bit architecture, what will happen is that a space of memory is addressable with 64 bit. As a consequence we can imagine that the addressable memory is space has $2^{64} - 1$ memory locations available.

  And each application sees, these much space available for its execution. But however large the space can be in a machine, it will never match the aforementioned size. It could potentially for one task, but in the case a machine is hundreds of tasks and each of them wants to use that much memory, there is no way that the hardware can provide enough physical memory to satisfy all of them.

  To counteract this problem, it is common practice to schedule the memory as well, because you take advantage of the fact that an application CAN use $2^{64} - 1$ memory locations, but at a given time it uses a tiny portion of these locations. It is only that tiny portion of memory locations that needs to be made available to the running task.

  In this scenario, the OS makes it possible to accommodate within the physical memory of the computer these small slices of the available space that the application uses. So somehow it operates as a resource manager for the memory as well.

  3. **File Management** File Management
  4. **Networking**, **Device Drivers**, **Graphical Interface** Networking Device Drivers Graphical Interface

  The important thing is that all of these resources, like the processor, the memory, the drivers etc..., are shared between all the tasks. All these resource managers have to be distributed among all the spectrum of tasks in such a way that the tasks behave properly, i.e. if you do not provide frequently enough these resources they would not be able to deliver the result on time (the OS manages this problem on its own).

In the case we decide to look at the Operating System as a Resource Manager, we need to think of a structure for the OS that makes this resource management effective, effective in the sense that we believe it is the most relevant for our specific range of application.
The way OSs handles devices, interrupt, etc. can be very different (and optimized in very different ways) depending on the type of application one is looking at. However, the type of optimizations we are interested in are those that allow our application to have time-limited execution.

Real-Time Systems

A **Real-Time application** is an application of which the time when a result is produced matters. Real-Time application
In particular:

- a correct result produced too late is equivalent to a wrong result, or to no result.
- it is characterized by temporal constraints that have to be respected.

Example 1: Mobile vehicle

Let us consider a mobile vehicle with a software module that

1. Detects obstacles
2. Computes a new trajectory to avoid them
3. Computes the commands for engine, brakes,...
4. Sends the commands

If you decide to steer to the left or to the right there is a limited amount of time in which the operation has to be carried out. Hence if one can find an extremely effective strategy for steering the wheels but the strategy amounts to setting the values for the motors after one second, it is completely useless, since the vehicle is most likely to crash.
Hence a time violation in executing a task is a critical problem: it means that the developed application is useless and also dangerous.
But then, what is a reasonable time frame for completing the steering operation?
Depends on the speed in which the vehicle is traveling. But no matters if the vehicle is traveling at high or low speed the timing constraint is there, and if it is violated, the vehicle will eventually crash against the obstacle.
As a consequence: when a constraint is set, that constraint needs to be respected. And this is one of the core concept of Real Time.

Hence, a Real-Time is not necessarily synonym of fast execution, but rather of **predictable** execution. predictable
Real time computing has much more to do with predictability than of being quick.
Some examples of temporal contraints are:

- The program must react to external events in a predictable time
- The program must repeat a given activity at a precise rate
- The program must end an activity before a specified time

In this case, we can clearly notice that the temporal constraints can be either one shot events or periodic events, but in both cases, a common characteristic, there is a need of being predictability.
Temporal constraints are modeled using the concept of **deadline**. deadline
Please notice that a Real-Time system is not just a *fast system*, because the speed is always relative to a specific environment, i.e. the steering commands temporal constraint is set by the velocity of the vehicle.
Running faster is good, but does not guarantee the correct behavior. In fact, it is far more valuable to that temporal constraints are always respected; in other terms Real time systems prefer to run fast enough to respect the deadlines, to be reliable.

Hence, the type of analysis that is necessary to perform is not an analysis based on of average/typical cases but rather an analysis of worst case: I have to prove that even in the worst-case scenario, there is not deadline violation.

This predictability creates a wide gap between what a Real Time system is and what a general purpose system is, because general purpose systems are optimised for the average case, but a real time system only cares about the worst case. As a consequence, the way one designs a Real Time system is very different from the way a general purpose system is designed.
In fact:

- When one optimize for the average case, what one would look at is the number of times that an application completes a task every second, and this is called **Throughput**. Throughput
- When one have a worst case requirement, the notion of throughput is not relevant anymore, and the analysis focuses in every single instance the maximum delay will be bounded.

Let us introduce some notion and general terms that we will extensively using during the course

Definition 4: Algorithm

Logical procedure used to solve a problem

Definition 5: Program

Formal description of an algorithm, using a *programming language*

Definition 6: Process

Instance of a program (program in execution)

Definition 7: Thread

Flow of execution, something that is able to execute using your processor along with other threads. These threads can be part of the same program and they can be executed in parallel.

Definition 8: Task

Process or thread

Hence there are two different ways of sharing resources: one are threads in which your share computing resources and memory space, and processes in which you share computer resources but each of the processes has its own memory space.
Unfortunately, there is no common definition of a task: somebody use the terms with the same meaning as a thread and sometimes it is used with the same meaning as a process. In this class we will refer to threads.
Henceforth, when we talk about a task we will refer to a program that it is running and they share the same memory space with other programs.

Chapter 1

Basic Concepts

A task can be seen as a sequence of actions and a deadline must be associated to each one of them. We, therefore, are after is a definition of a formal model that identifies what these tasks or actions are and associate deadlines with them.

1.1 Real-Time Tasks

Definition 9: Real-Time Task (τ_i)

stream of jobs (or instances) $J_{i,k}$, or, in other terms, a sequence of activities that is activated periodically or aperiodically

Each job $J_{i,k} = (r_{i,k}, c_{i,k}, d_{i,k})$ is characterised by the following quantities:

- $r_{i,k}$ activation time — `activation time`
  It is the time at which a task becomes ready for execution; it is also referred as *request time* or *release time*.
- $c_{i,k}$ computation time — `computation time`
  Time necessary to the processor for executing the job without interruption.
- $d_{i,k}$ absolute deadline — `absolute deadline`
  time before which a job should be completed to avoid damage to the system.
- $f_{i,k}$ finishing time — `finishing time`
  The time at which a job finishes its execution
- $\rho_{i,k}$ response time — `response time`
  The time at which a job finishes its execution. Formally this quantity is the difference between the finishing time and the activation time.

$$\rho_{i,k} = f_{i,k} - r_{i,k}$$

Furthermore, since each task i is a sequence of jobs, we need to differentiate between them. That is why each job $J_{i,k}$ is uniquely identified by its task index i and the k-th activation of the i-th task. In addition, we will say that job $J_{i,k}$ respects its deadline if $f_{i,k} \leq d_{i,k}$.

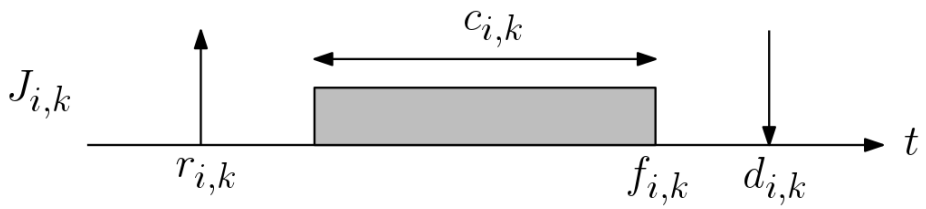


Figure 1.1: Graphical representation of Mathematical model of a Task

This mathematical definition of a job in a real-time task holds regardless of the nature of the task itself. In fact, we can identify three different types of tasks: Periodic tasks, Aperiodic Tasks and Sporadic Tasks. Each of them holds different properties and a different mathematical representation.

1.1.1 Periodic Tasks

Definition 10: Periodic Task

A periodic task $\tau_i = (C_i, D_i, T_i)$ is a stream of jobs $J_{i,k}$, with:

$$\begin{aligned} r_{i,k+1} &= r_{i,k} + T_i \\ d_{i,k} &= r_{i,k} + D_i \\ C_i &= \max_k \{c_{i,k}\} \end{aligned}$$

where:

- T_i **Period**
- D_i **Relative Deadline**
- C_i **Worst-Case Execution Time (WCET)**
- R_i **Worst-Case Response Time (WCRT)**

$$R_i = \max_k \{\rho_{i,k}\} = \max_k \{f_{i,k} - r_{i,k}\}$$

For the task to be correctly scheduled, it must be $R_i \leq D_i$

A periodic task has a regular structure (called **cycle**), in the sense that:

- it is activated periodically with a period of T_i
- it executes a computation
- when the computation terminates, it suspends waiting for the next period

Hence, its fundamental implementation can be represented as:

```
void *PeriodicTask(void *arg)
{
  <initialization>;
  <start periodic timer, period = T>;
  while (condition)
  {
    <read sensors>;
    <update outputs>;
    <update state variables>;
    <wait next activation>;
  }
}
```

1.1.2 Aperiodic Tasks

Definition 11: Aperiodic Task

Aperiodic tasks are not characterised by periodic arrivals, meaning that:

- A minimum interarrival time between activations does not exist
- Sometimes, aperiodic tasks do not have a particular structure

Aperiodic tasks can model tasks responding to events that occur rarely (e.g. a mode change) or tasks responding to events with irregular structure (e.g. bursts of packets from the network,...).

1.1.3 Sporadic Tasks

Sporadic tasks are aperiodic tasks characterised by a **Minimum Interarrival Time (MIT)** between jobs. In this sense they are similar to periodi tasks, but while a periodic task is activated by a periodic timer, a sporadic task is activated by an external event. (e.g. the arrival of a packet from the network)

Hence, its fundamental implementation can be represented as:

```
void *SporadicTask(void *arg)
{
    <initialization>;
    while (condition)
    {
        <computation>;
        <wait events>;
    }
}
```

Formally:

Definition 12: Sporadic Task

A sporadic task $\tau_i = (C_i, D_i, T_i)$ is a stream of jobs $J_{i,k}$, with:

$$r_{i,k+1} \geq r_{i,k} + T_i$$
$$d_{i,k+1} = r_{i,k} + D_i$$
$$C_i = \max_k \{c_{i,k}\}$$

where:

- T_i **Minimum Interarrival Time (MIT)**
- D_i **Relative Deadline**
- C_i **Worst-Case Execution Time (WCET)**
- R_i **Worst-Case Response Time (WCRT)**

Minimum Interarrival Time (MIT)

Relative Deadline

Worst-Case Execution Time (WCET)

Worst-Case Response Time (WCRT)

$$R_i = \max_k \{\rho_{i,k}\} = \max_k \{f_{i,k} - r_{i,k}\}$$

For the task to be correctly scheduled, it must be $R_i \leq D_i$.

1.2 Task Criticality

A deadline is said to be *hard* if a deadline miss causes a critical failure in the system, whereas a task is said to be a **hard real-time task** if all its deadlines are hard, which means that all the deadlines must be guaranteed before starting the task, i.e.

hard real-time task

$$\forall j, \rho_{i,j} \leq D_i \quad \Rightarrow \quad R_i \leq D_i$$

Example 2: Hard Real-Time Task

The controller of a mobile robot, must detect obstacles and react within a time dependent on the robot speed, otherwise the robot will crash into the obstacles

A deadline is said to be *soft* if a deadlien miss causes a degradation in the **Quality of Service (QoS)**, but is not a catastrophic event, whereas a task is said to be a **soft real-time task** if it has soft deadlines.

Quality of Service (QoS)

soft real-time task

In other terms, some deadlines can be missed without compromising the correctess of the system, but the number of missed deadlines must be kept under control, because the *quality* of the results depend on the number of missed deadlines.

Unline the hard real-time task, soft real-time tasks can be difficult to characterize, particularly:

- What's the tradeoff between *non compromising the system correctness* and *not considering missed deadlines*?
- Moreover, some way to express the QoS experienced by a soft real-time task is needed

Examples of QoS definitions could be

- no more than X consecutive deadlines can be missed
- no more that X deadlines in an interval of time T can be missed
- the **deadline miss probability** must be less than a specified value, i.e. deadline miss probability

$$P\{f_{i,j} > d_{i,j}\} \leq R_{max}$$

- the **deadline miss ratio** must be less than a specified value, i.e. deadline miss ratio

$$\frac{\text{number of missed deadlines}}{\text{total number of deadlines}} \leq R_{max}$$

- the maximum **tardiness** must be less than a specified value, i.e. tardiness

$$\frac{R_i}{D_i} < L$$

- ...

Example 3: Audio and Video players

Assuming a framerate of 25 fps, which imply a frame period of 40 ms, if a frame is played a little bit too late, the user might even be unable to notice any degration in the QoS, however, skipped frames can be disturbing.
In fact missing a lot of frames by 5 ms can be better than missing only a few frames by 40 ms.

Example 4: Robotic Systems

Some actuations can be delayed with little consequences on the control quality.

In any case, soft real-time constraints does not mean no guarantee on dealines, given that tasks can have variable execution times between different jobs.
These execution times might depend on different factors:

- Input data
- HW issues (cache effects, pipeline stalls, ...)
- The internal state of the task
- ...

1.3 Schedulability analysis

Schedulability analysis tries to answer the question: Given a task set $\mathcal{T}$, how can we guarantee if it is schedulable or not?

1.3.1 Simulating the hyperperiod

The first possibility is to simulate the system to check that no deadline is missed. The execution time of every job is set equal to the WCET of the corresponding task.

In the case of periodic tasks with no offsets it is sufficient to simulate the schedule until the **hyperperiod** (H = $lcm\{T_i\}$). hyperperiod
In the case of offsets $\phi_i = r_{i,0}$ it is sufficient to simulate until $2H + \phi_{max}$.
If tasks periods are prime numbers the hyperperiod can be very large!

In the case of sporadic tasks, we can assume them to arrive at the highest possible rate, so we fall back to the case of periodic tasks with no offsets.

1.3.2 (Worst-Case) Response Time Analysis

According to the methods proposed by Audsley et al., the longest response time R_i of a periodic task τ_i is computed, at the critical instant, as the sum of its computation time and the interference I_i of the higher priority tasks:

$$R_i = C_i + I_i$$

where:

$$I_i = \sum_{j=1}^{i-1} \left\lceil \frac{R_i}{T_j} \right\rceil C_j$$

Hence,

$$R_i = C_i + \sum_{j=1}^{i-1} \left\lceil \frac{R_i}{T_j} \right\rceil C_j \tag{1.1}$$

Definition 13: Critical instant

The Critical instant for task τ_i occurs when job $J_{i,j}$ is released at the same time with a job in every high priority task

It is straighforward to notice that if all the offsets of the task set are 0, the first job of every task is released at the **critical instant**. critical instant

A job $J_{i,j}$ released at the critical instant experiences the maximum response time for τ_i:

$$\forall k, \quad \rho_{i,j} \geq \rho_{i,k}$$

No simple solution exists for this equation since R_i appears on both sides of the equation. Thus, the worst-case response time of task τ_i is given by the smallest value of R_i that satisfies equation 1.1. Notice, however, that only a subset of points in the interval $[0, D_i]$ need to be examined for feasibility. In fact, the interference on τ_i only increases when there is a release of a higher-priority task.

To simplify the notation, let $R_i^{(k)}$ be the k-th estimate of R_i and let $I_i^{(k)}$ be the interference on task τ_i in the interval $[0, R_i^{(k)}]$

$$I_i^{(k)} = \sum_{j=1}^{i-1} \left\lceil \frac{R_i^{(k)}}{T_j} \right\rceil C_j \tag{1.2}$$

Then the calculation of R_i is performed as follows:

1. Iteration starts with $R_i^{(0)} = \sum_{j=1}^{i} C_j$, which is the first point in time that τ_i could possibly complete
2. The actual interference I_i^k in the interval $[0, R_i^{(k)}]$ is computed by equation 1.2
3. If $I_i^{(k)} + C_i = R_i^{(k)}$, then $R_i^{(k)}$ is the actual worst-case response time of task τ_i; that is, $R_i = R_i^{(k)}$. Otherwise, the next estimate is given by

   $$R_i^{(k+1)} = I_i^{(k)} + C_i$$

   and the iteration continues from step 2.

Once R_i is calculated, the feasibility of task τ_i is guaranteed if and only if $R_i \leq D_i$.

The response time analysis is an efficient algorithm: in the worst case, the number of steps N for the algorithm to converge is exponential and it depends on the total number of jobs of higher priority tasks in the interval $[0, D_i]$:

$$N \propto \sum_{h=1}^{i-1} \left\lceil \frac{D_h}{T_h} \right\rceil$$

If s is the minimum granularity of the time, then in the worst case $N = \frac{D_i}{s}$. However, such worst case is very rare, usually the number of steps is low.

1.3.3 Processor Demand Analysis

Another necessary and sufficient test for checking the schedulability of fixed priority systems with constrained deadlines was proposed by Lehoczky, Sha and Ding. The test is based on the concept of Level-i workload, defined as follows

Definition 14: Level-i workload

The Level-i workload $W_i(t)$ is the cumulative computation time requested in the interval $(0, t]$ by task τ_i and all the tasks with priority higher than p_i

The basic idea is very simple: in any interval, the computation demanded by all tasks in the set must never exceed the available time.
The problem is: how to compute the time demanded by a tast set $\mathcal{T}$?
Since we have to look only at jobs released at the critical instant, we can consider all offsets equal to zero and only consider the first job of each task...

Definition 15: Processor Demand

Given an interval $[t_1, t_2]$, let $\mathcal{J}_{t_1,t_2}$ be the set of jobs started after t_1 and with deadline lower than or equal to t_2:

$$\mathcal{J}_{t_1,t_2} = \{J_{i,j} : r_{i,j} \geq t_1 \wedge d_{i,j} \leq t_2\}$$

The processor demand in $[t_1, t_2]$ is defined as:

$$W(t_1, t_2) = \sum_{J_{i,j} \in \mathcal{J}_{t_1,t_2}} c_{i,j}$$

Worst case: use C_i instead of $c_{i,j}$

Guaranteeing a task set $\mathcal{T}$ based on $W(t_1, t_2)$ can take a long time.
In fact, it must hold

$$\forall (t_1, t_2) \quad W(t_1, t_2) \leq t_2 - t_1$$

This means that the test requires to check all the (t_1, t_2) combinations in a hyperperiod.
However, we only need to check the first job of every task τ_i.

The quantity $W_i(t_1, t_2)$ is the time demanded in $[t_1, t_2]$ by all tasks τ_j with $p_j \geq p_i$ ($\Rightarrow j \leq i$)

We can consider only $W_i(0, t)$.
For task τ_i only check $W_i(0, t)$ for $0 \leq t \leq D_i$.
Change $\forall$ into $\exists$: consider worst case for $W_i()$
The number of jobs in $[0, t]$ is $\left\lfloor \frac{t}{T_i} \right\rfloor$
Use $\lceil\rceil$ instead

We already have hints about computing an upper bound for $W_i(0, t)$...

$$W_i(0, t) = C_i + \sum_{h=1}^{i-1} \left\lceil \frac{t}{T_h} \right\rceil C_h$$

Task τ_i is schedulable if and only if $\exists t : 0 \leq t \leq D_i \wedge W_i(0, t) \leq t$.
A task set $\mathcal{T}$ is schedulable if and only if

$$\forall \tau_i \in \mathcal{T}, \quad \exists t : 0 \leq t \leq D_i \wedge W_i(0, t) \leq t$$

Sometimes, different notations in literature:

$$W_i(0, t) \rightarrow W_i(t) = \sum_{h=1}^{i} \left\lceil \frac{t}{T_h} \right\rceil C_h$$

This is equivalent, because $0 \leq t \leq T_i$.
Someone defines

$$L_i(t_1, t_2) = \frac{W_i(t_1, t_2)}{t_2 - t_1}$$

$$L_i = \min_{0 \le t \le D_i} L_i(0,t) \quad ; \quad L = \max_{\tau_i \in \mathcal{T}} L_i$$

The guarantee tests then becomes:

- Task τ_i is schedulable iff $L_i \le 1$
- $\mathcal{T}$ is schedulable iff $L \le 1$

The test might still be long (need to check many values of $L(0,t)$ to find the minimum)...
The number of points to check for computing W_i or L_i can be reduced:

$$S_i = \left\{ k\,T_h | h \le i; 1 \le k \le \left\lfloor \frac{T_i}{T_h} \right\rfloor \right\}$$

multiples of T_h for $h \le i$

$$L_i = \min_{t \in S_i} L_i(0,t)$$

1.3.4 Processor Utilization Factor test

The feasibility of a task set with contrained deadlines could be guaranteed using the utilization based test, by reducing tasks' periods to relative deadlines:

$$U_{lub} = \sum_{i=1}^{n} \frac{C_i}{D_i} \le n(2^{1/n} - 1)$$

However, such a test would be quite pessimistic, since the workload on the processor would be overestimated.
For this reason this test is **sufficient but not necessary**.

Nonetheless, in many cases it is useful to have a very simple test to see if a task set is schedulable. This sufficient test is based on the **Utilisation bound**.

Utilisation bound

Definition 16: Utilisation Least Upper Bound

The utilisation least upper bound for a scheduling algorithm $\mathcal{A}$ is the smallest possible utilisation U_{lub} such that, for any task set $\mathcal{T}$, if the task set's utilisation U is not greater than U_{lub} ($U \le U_{lub}$), then the task set is schedulable by algorithm $\mathcal{A}$

In other terms, we can consider that each task uses the processor for a fraction of time

$$U_i = \frac{C_i}{T_i}$$

The total processor utilisation is

$$U = \sum_i \frac{C_i}{T_i}$$

which we will consider as a measure of the processor's load.

Given these definition, the necessary condition for the schedulability of a task set is:

- If $U > 1$ the task set is surely not schedulable
- If $U \le U_{lub}$, the task set is schedulable
- If $U_{lub} < U \le 1$ the task set may or may not be schedulable

Ideally a value of $U_{lub} = 1$ would be optimal.

In general, given that the tasks might not always have relative deadline equals to the period the formulation of the total processor utilisation considers the relative deadline:

$$U' = \sum_{i=1}^{n} \frac{C_i}{D_i}$$

This approach considers the worst case for a task... hence if the task set is guaranteed using the relative deadlines, it must hold that the test holds even when considering the period.

The bound is very pessimistic: most of the times, a task set with $U > U_{lub}$ is schedulable. A particular case is when tasks have periods that are harmonic.

Definition 17: Harmonic task set

A task set is harmonic if, for every two tasks τ_i, τ_j either T_i is multiple of T_j or T_j is multiple of T_i

For a harmonic task set, the utilisation bound is $U_{lub} = 1$. (Foreshadowing: Rate Monotonic is an optimal algorithm for harmonic task sets)

1.3.5 Examples of schedulability analysis

1.3.5.1 Example 1

Consider a task set of three periodic tasks with deadline equal to period:

$$\tau_1 = (20, 100) \qquad \tau_2 = (40, 150) \qquad \tau_3 = (100, 350)$$

Now let us consider the schedulability of the task set using the three methods introduced before:

Example 5: Processor Utilization Factor test

First, let us compute U_{lub} for the task set of three tasks with $n = 3$

$$U_{lub} = n(2^{1/n} - 1) = 3(2^{1/3} - 1) \approx 0.77976315$$

Then let us compute the Utilisation for the three tasks:

$$U = \sum_{i=1}^{n} \frac{C_i}{T_i} = \frac{20}{100} + \frac{40}{150} + \frac{100}{350} = 0.752380952$$

Hence, the sufficient test states that since $U \leq U_{lub}$ the task set is schedulable.

Example 6: Processor Demand Analysis

Let us choose a set of scheduling points for the analysis.

- Task τ_1.

  $$S_1 = \{100\}$$

  For each scheduling point in S_1 find, if any, point for which $W_1(t) \leq D_1$

  $$W_1(0, 100) = \sum_{h=1}^{1} \left\lceil \frac{t}{T_h} \right\rceil C_h = C_1 = 20 \leq 100$$

  Since there exists at least a scheduling point where the relationship is satisfied, τ_1 is schedulable
- Task τ_2.

  $$S_2 = \{100, 150\}$$

  $$W_2(0, 100) = \sum_{h=1}^{2} \left\lceil \frac{t}{T_h} \right\rceil C_h = 20\frac{100}{100} + 40 \left\lceil \frac{100}{150} \right\rceil = 20 + 40 = 60 \leq 100$$

  Since there exists at least a scheduling point where the relationship is satisfied, τ_2 is schedulable
- Task τ_3.

  $$S_3 = \{100, 150, 200, 300, 350\}$$

$$W_3(0, 100) = 20 + 40 + 100 = 160 > 100$$
$$W_3(0, 150) = 2 \cdot 20 + 40 + 100 = 180 > 150$$
$$W_3(0, 200) = 2 \cdot 20 + 2 \cdot 40 + 100 = 220 > 200$$
$$W_3(0, 300) = 3 \cdot 20 + 2 \cdot 40 + 100 = 240 \leq 300$$

Since there exists at least a scheduling point where the relationship is satisfied, τ_3 is schedulable

Example 7: Response Time Analysis

- Task τ_1.

$$R_1^{(0)} = C_1 = 20$$
$$R_1^{(1)} = C_1 = R_1^{(0)}$$

Since $R_1 < D_1$, the task is schedulable.

- Task τ_2.

$$R_2^{(0)} = C_2 = 40$$
$$R_2^{(1)} = 40 + \left\lceil \frac{40}{100} \right\rceil 20 = 60$$
$$R_2^{(2)} = 40 + \left\lceil \frac{60}{100} \right\rceil 20 = R_2^{(1)}$$

Since $R_2 \leq D_2$ the task is schedulable.

- Task τ_3.

$$R_3^{(0)} = C_3 = 100$$
$$R_3^{(1)} = 100 + \left\lceil \frac{100}{100} \right\rceil 20 + \left\lceil \frac{100}{150} \right\rceil 40 = 160$$
$$R_3^{(2)} = 100 + \left\lceil \frac{160}{100} \right\rceil 20 + \left\lceil \frac{160}{150} \right\rceil 40 = 220$$
$$R_3^{(3)} = 100 + \left\lceil \frac{220}{100} \right\rceil 20 + \left\lceil \frac{220}{150} \right\rceil 40 = 240$$
$$R_3^{(4)} = 100 + \left\lceil \frac{240}{100} \right\rceil 20 + \left\lceil \frac{240}{150} \right\rceil 40 = R_3^{(3)}$$

Since $R_3 \leq D_3$ the task is schedulable.

1.3.5.2 Example 2

Consider a task set of three periodic tasks with deadline equal to period:

$$\tau_1 = (40, 100) \qquad \tau_2 = (40, 150) \qquad \tau_3 = (100, 350)$$

Now let us consider the schedulability of the task set using the three methods introduced before:

Example 8: Processor Utilization Factor test

First, let us compute U_{lub} for the task set of three tasks with $n = 3$

$$U_{lub} = n(2^{1/n} - 1) = 3(2^{1/3} - 1) \approx 0.77976315$$

Then let us compute the Utilisation for the three tasks:

$$U = \sum_{i=1}^{n} \frac{C_i}{T_i} = \frac{40}{100} + \frac{40}{150} + \frac{100}{350} = 0.952380952$$

Since $U_{lub} < U < 1$ we cannot determined whether the task set is schedulable using the Processor Utilitzation Factor test.

Example 9: Processor Demand Analysis

Let us choose a set of scheduling points for the analysis.

- Task τ_1.

  $$S_1 = \{100\}$$

  For each scheduling point in S_1 find, if any, point for which $W_1(t) \leq D_1$

  $$W_1(0, 100) = \sum_{h=1}^{1} \left\lceil \frac{t}{T_h} \right\rceil C_h = C_1 = 40 \leq 100$$

  Since there exists at least a scheduling point where the relationship is satisfied, τ_1 is schedulable
- Task τ_2.

  $$S_2 = \{100, 150\}$$

  $$W_2(0, 100) = \sum_{h=1}^{2} \left\lceil \frac{t}{T_h} \right\rceil C_h = 40\frac{100}{100} + 40 \left\lceil \frac{100}{150} \right\rceil = 40 + 40 = 80 \leq 100$$

  Since there exists at least a scheduling point where the relationship is satisfied, τ_2 is schedulable
- Task τ_3.

  $$S_3 = \{100, 150, 200, 300, 350\}$$

  $$\begin{aligned} W_3(0, 100) &= 40 + 40 + 100 = 180 > 100 \\ W_3(0, 150) &= 2 \cdot 40 + 40 + 100 = 220 > 150 \\ W_3(0, 200) &= 2 \cdot 40 + 2 \cdot 40 + 100 = 260 > 200 \\ W_3(0, 300) &= 3 \cdot 40 + 2 \cdot 40 + 100 = 300 \leq 300 \end{aligned}$$

  Since there exists at least a scheduling point where the relationship is satisfied, τ_3 is schedulable

Example 10: Response Time Analysis

- Task τ_1.

  $$\begin{aligned} R_1^{(0)} &= C_1 = 40 \\ R_1^{(1)} &= C_1 = R_1^{(0)} \end{aligned}$$

  Since $R_1 < D_1$, the task is schedulable.

- Task τ_2.

$$R_2^{(0)} = C_2 = 40$$
$$R_2^{(1)} = 40 + \left\lceil \frac{40}{100} \right\rceil 40 = 80$$
$$R_2^{(2)} = 40 + \left\lceil \frac{80}{100} \right\rceil 40 = R_2^{(1)}$$

  Since $R_2 \leq D_2$ the task is schedulable.
- Task τ_3.

$$R_3^{(0)} = C_3 = 100$$
$$R_3^{(1)} = 100 + \left\lceil \frac{100}{100} \right\rceil 40 + \left\lceil \frac{100}{150} \right\rceil 40 = 180$$
$$R_3^{(2)} = 100 + \left\lceil \frac{180}{100} \right\rceil 40 + \left\lceil \frac{180}{150} \right\rceil 40 = 260$$
$$R_3^{(3)} = 100 + \left\lceil \frac{260}{100} \right\rceil 40 + \left\lceil \frac{260}{150} \right\rceil 40 = 300$$
$$R_3^{(4)} = 100 + \left\lceil \frac{300}{100} \right\rceil 40 + \left\lceil \frac{300}{150} \right\rceil 40 = R_3^{(3)}$$

  Since $R_3 \leq D_3$ the task is schedulable.

Real-Time Scheduling

Chapter 2

Periodic Task Scheduling

The term task is used to indicate a schedulable entity (either a process or a thread), in particular:

- A thread represents a flow of execution (it executes with shared resources, multi thread within the same process)
- A process represents a flow of execution + private resources (it executes with its own resources), such as address space, file table, ...

Tasks do not run on bare hardware, but then how can multiple tasks execute on one single CPU? The OS kernel is a piece of the operating system that takes care of multi-programming and somehow it is able to create the illusion that each CPU/processor has its own space, whereas in fact it is sharing the same resources with other processes.
In the end the kernel provides the mechanism that enable multiple tasks to execute in parallel; in a sense tasks have the illusion of executing concurrently on a dedicated CPU per task.

On this regard, with the term concurrency we refer to the simultaneous execution of multiple threads/processes in the same PC.
Concurrency is implemented by multiplexing tasks on the same CPU. Tasks are alternated on a real CPU and the task scheduler decides which task executes at a given instant in time. In other terms, in order to implement the concurrency mechanism it is necessary to introduce this new component (i.e. the task scheduler), since it makes sure that the time of your pc is shared between the different processes or tasks that compete for the reosurces at that time.

Tasks are associated to temporal constraints (a.k.a. deadlines), hence the scheduler must allocate the CPU to tasks so that their deadlines are respected.

2.1 Real Time Scheduling

Definition 18: Scheduler

A scheduler generates a schedule from a set of tasks

1. In the case of Unicore processor system (UP) (simpler definition), a schedule $\sigma(t)$ is a function mapping time t into an executing task.

   $$\sigma : t \to \mathcal{T} \cup \tau_{idle}$$

   where $\mathcal{T}$ is the taskset and τ_{idle} is the idle task

2. For a Symmetric Multipprocessor System (SMP) (m CPUs), $\sigma(t)$ can be extended to map t in vectors $\tau \in (\mathcal{T} \cup \tau_{idle})^m$

Hence a scheduler is responsible for selecting the task to execute at time t.

Definition 19: Scheduling algorithm

Algorithm used to select for each time instant t a task to be executed on a CPU among the ready task

Given a task set $\mathcal{T}$, a scheduling algorithm $\mathcal{A}$ generates the schedule $\sigma_{\mathcal{A}}(t)$.
A task set is schedulable by an algorithm $\mathcal{A}$ if $\sigma_{\mathcal{A}}$ does not contain missed deadlines.
To verify that no missed deadlines occur, a **Schedulability test** checks if $\mathcal{T}$ is schedulable by $\mathcal{A}$.

Schedulability test

2.2 Cyclic Executive Scheduling

Timeline Scheduling, also known as **Cyclic Executive Scheduling**, is one of the most used approaches to handle periodic tasks in defense military systems and traffic control systems.
The methods consists of dividing the tmeporal axis into slots of equal length, in which one or more tasks can be allocated for execution, in such a way to respect the frequencies derived from the application requirements. A timer synchronizes the activation of the tasks at the beginning of each time slot.

Timeline Scheduling

Cyclic Executive Scheduling

Cyclic Executing Scheduling is a **static scheduling algorithm** where **jobs are not preemptable** (i.e. A scheduled job executes until termination).
The slots are statically allocated to the tasks using a **scheduling table**.
In this Scheduling algorithm two quantities are considered:

static scheduling algorithm

scheduling table

- **Major Cycle**: least common multiple of all the tasks' periods (a.k.a. **hyperperiod**)
- **Minor Cycle**: greatest common divisor of all the tasks' periods

Major Cycle

hyperperiod

The period timer fires every Minor Cycle Δ.
Hence the implementation of the scheduling algorithm performs as follow:

Minor Cycle

1. The periodic timer fires every minor cycle
2. Read the scheduling table and execute the appropriate tasks
3. Sleep until next minor cycle

The main advantage of timeline scheduling is its simplicity. The method can be implemented by programming a timer to interrupt with a period equal to the minor cycle and by writing a main program that calls the tasks in the order given in the major cycle, inserting a time synchronization point at the beginning of each minor cycle. Since the task sequence is not decided by a scheduling algorithm in the kernel, but it is triggered by the calls made by the main program, there are no context switches, so the runtime overhead is very low. Moreover, the sequence of tasks in the schedule is always the same, can be easily visualized, and it is not affected by jitter (i.e., task start times and response times are not subject to large variations).

In spite of these advantages, timeline scheduling has some problems. For example, it is very fragile during overload conditions. If a task does not terminate at the minor cycle boundary, it can either be continued or aborted. In both cases, however, the system may run into a critical situation. In fact, if the failing task is left in execution, it can cause a domino effect on the other tasks, breaking the entire schedule (timeline break). On the other hand, if the failing task is aborted while updating some shared data, the system may be left in an inconsistent state, jeopardizing the correct system behavior.
Another big problem of the timeline scheduling technique is its sensitivity to application changes. If updating a task requires an increase of its computation time or its activation frequency, the entire scheduling sequence may need to be reconstructed from scratch.
Finally, another limitation of the timeline scheduling is that it is difficult to handle aperiodic activities efficiently without changing the task sequence. The problems outlined above can be solved by using priority-based scheduling algorithms.

2.3 Fixed Priority Scheduling

Fixed Priority Scheduling is a very simple preemptive scheduling algorithm.

To each task τ_i is assigned a fixed priority p_i as an integer number: the higher the number the higher the priority. In the research literature sometimes, authoers use the opposite convention: the lowest the number, the highest the priority.

The active task with the highest priority is scheduled.

Fixed Priority Scheduling has the following priority:

- The response time of the task with the highest priority is minimum and equal to its WCET
- The reposnse time of the other tasks depends on the interference of the higher priority tasks
- The priority assignment may influence the schedulability of a task set
  Problem: how to assign tasks' priorities so that a task set is schedulable?

There are two main approaches to assigning priorities to the task set:

- **Schedulability**, i.e. find the priority assignment that makes all tasks schedulable — Schedulability
- **Response time (optimization)**, i.e. find the priority assignment that minimise the response time of a subset of tasks — Response time (optimization)

By now we consider the first objective only, hence we will investigate the **optimal priority assignment (Opt)**. — optimal priority assignment (Opt)

2.3.1 Rate Monotonic Scheduling

The Rate Monotonic (RM) scheduling algorithm is a simple rule that assigns priorities to tasks according to their request rates. Specifically, tasks with higher request rates (that is, with shorter periods) will have higher priorities. Since periods are constant, RM is a fixed-priority assignment: a priority p_i is assigned to the task before execution and does not change over time. Moreover, RM is intrinsically preemptive: the currently executing task is preempted by a newly arrived task with a shorter period.

In 1973, Liu and Lyland showed that RM is **optimal** among all fixed-priority assignments (with deadline equals to the period and offset equal to 0) in the sense that no other fixed-priority algorithms can schedule a task set that cannot be scheduled by RM.
In addition, RM is an optimal algorithm for harmonic task sets. This holds also for sporadic tasks.

2.3.2 Deadline Monotonic Scheduling

The Dealine Monotonic (DM) priority assignemt weakens the *periodi equals dealine* contraint within a static priority scheduling scheme. This algorithm was first proposed in 1982 by Leung and Whitehead as an extension of Rate Monotonic, where tasks can have relative dealines less than or equal to their period (i.e. *constrained deadlines*).

According to the DM algorithm, each task is assigned a fixed priority p_i inversely proportional to its relative deadline D_i. Thus, at any instant, the task with the shorter relative deadline is executed. Since relative deadlines are constant, DM is a static priority assignment. As RM, DM is normally used in a fully preemptive mode, that is the currently executing task is preempted by a newly arrived task with shorter relative deadline.

The DM priority assignment is **optimal**, meaning that, if a task set is schedulable by some fixed priority assignment (with deadline different from the period and offset equal to 0), then it is also schedulable by DM.
This holds also for sporadic tasks.

2.4 Dynamic Priority Scheduling

RM and DM are optimal fixed priority assignments. Maybe we can improve schedulability by using **dinamic priorities**? Assumption: priorities change from job to job (a job $J_{i,j}$ always has the same priority $p_{h,k}$) — dinamic priorities

2.4.1 Earliest Deadline First (EDF)

The Earliest Deadline First (EDF) algorithm is a dynamic scheduling rule that selects tasks according to their absolute deadlines. Specifically, tasks with earlier deadlines will be executed at higher priorities. Since the absolute deadline of a periodic task depends on the current j-th instance as

$$d_{i,j} = (j-1)T_i + D_i$$

EDF is a dynamic priority assignment. Moreover, it is typically executed in preemptive mode, thus the currently executing task is preempted whenever another periodic instance with realier deadline becomes active.

Note that EDF does not make any specific assumption on the periodicity of the tasks; hence, it can be used for scheduling periodc as well as aperiodic and sporadic tasks.

The most important benefit of the EDF scheduling is that if we consider a task set $\mathcal{T}$ of periodic tasks with deadline equal to period, we know that a necessary condition for schedulability is that the sum of utilization is to be lower than 1

$$U = \sum_i \frac{C_i}{T_i} \leq 1$$

On top of this the processor utilization factor test has an upper bound U_{lub} that measures how effective is the algorithm, which is a sufficient condition.

The nice thing about EDF scheduling is that the least upper bound $U_{lub} = 1$. Hence, the processor utilization factor test is a sufficient and necessary condition for the schedulability of a task set. In other terms, using EDF achieves the full utilization of the processor.

Notice that this does not mean that RM is not optimal. RM is optimal among the task set of periodic tasks with deadline equals period and static priority assignment.
In this case, EDF is optimal among the scheduling algorithms with dynamic priority assignment and deadline equal to the period.

In case the deadline is different from the period, the schedulability test of EDF needs to use Response Time Analysis and Processor demand analysis with a slight adjustment, according to the choice of deadlines.

There are some cases when a task set not schedulable with RM can be scheduled using EDF. Consider the following task set of periodic tasks with deadline equals to the period:

$$\begin{cases} \tau_1 = (3, 8, 8) \\ \tau_2 = (6, 11, 11) \end{cases}$$

The utilisation of this task set is $U = 0.92$

As seen in figure 2.1, task τ_2 misses its deadline.

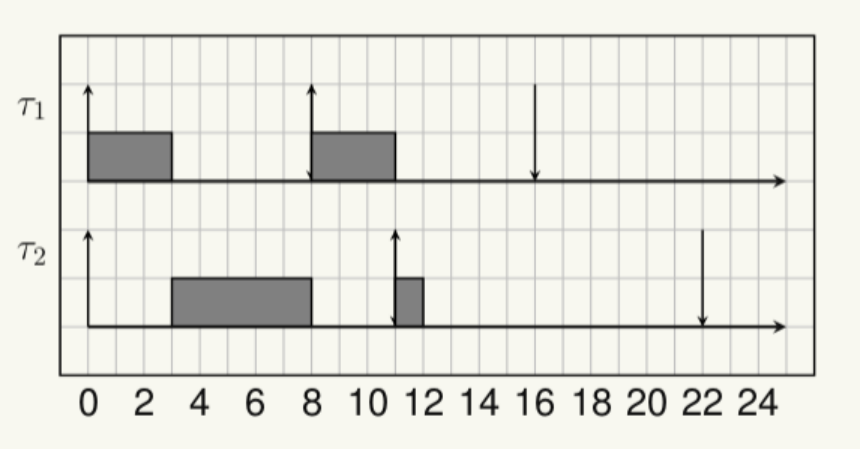


Figure 2.1

Using EDF the task set is schedulable as seen in figure 2.2

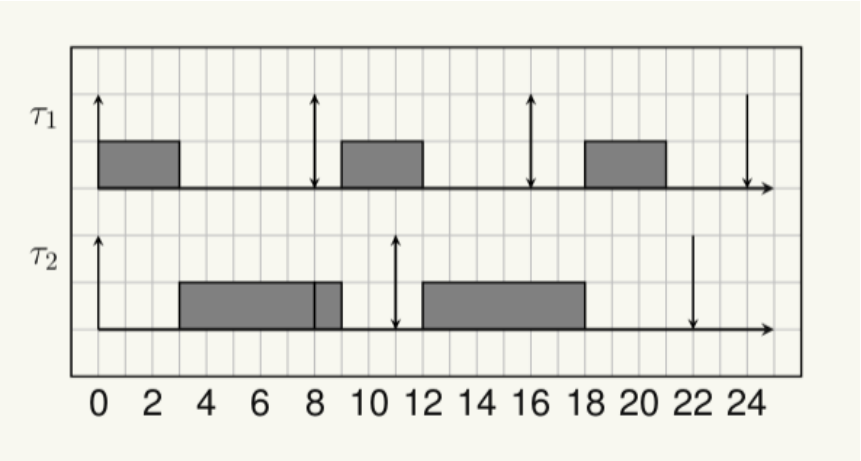


Figure 2.2

EDF from a theoretically point of view solves all our problems in terms of schedulability, but most real-time operating system utilizes fixed priority assignments.
This is because priorities are useful for scheduling analysis and because the priority allows the developer to quantify the relative importance of a task.

Chapter 3

Aperiodic Servers

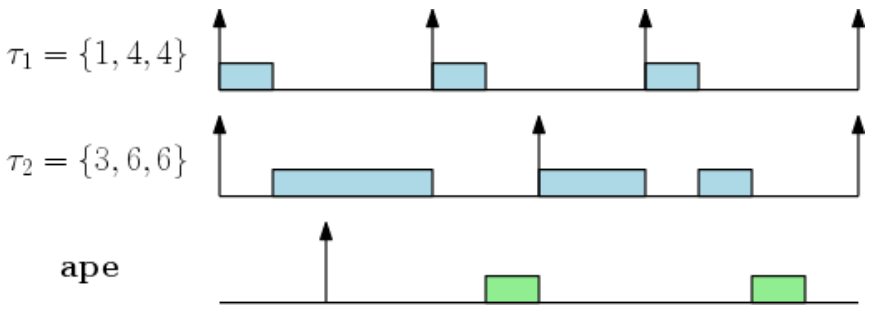
The scheduling algorithms treated in the previous chapter deals with homogeneous sets of tasks, where all computational activities are periodic. Many real-time control applications, however, require both aperiodic and periodic processes, which may also differ for their criticality. Tipically, periodic tasks are time-driven and execute critical control activities with hard timing contraints aimed at guaranteeing regular activation rates. Aperiodic tasks are usually event-driven and may have hard, soft, or non real-time requirements depending on the specific applications.

When dealing with hybrid task sets, the main objective of the kernel is to guarantee the schedulability of all critical tasks in worst-case conditions and provide good average response times for soft and non-real-time activities. Off-line guarantee of event-driven aperiodic tasks with critical timing contraints can be done only by making proper assumptions on the environment; that is, by assuming a maximum arrival rate for each critical event. This implies that aperiodic tasks associated with critical events are characterized by a minimum interarrival time between consecutive instances, which vounds the aperiodic load. Aperiodic tasks characterized by a minimum interarrival time are called sporadic. They are guaranteed under peak-load situations by assuming their maximum arrival rate.

3.1 Background Execution

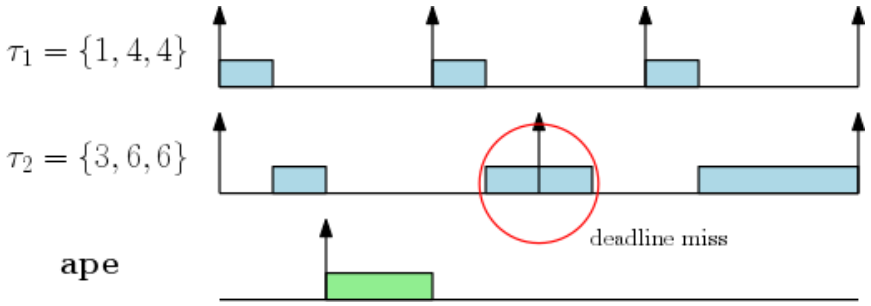
The simplest method to handle a set of soft aperiodic activities in the presence of periodic tasks is to schedule them in background; that is, when there are not periodic instances ready to execute. The major problem with this technique is that, for high periodic loads, the response time of aperiodic requests can be too long for certain applications. For this reason, background scheduling can be adopted only when the aperiodic activities do not have stringest timing constraints and the periodic load is not high.

The major advantage of background scheduling is its simplicity. In general, only two queues are needed to implement the scheduling mechanism: one (with a higher priority) dedicated to periodic tasks and the other (with a lower priority) reserved for aperiodic requests. The two queueing strategies are independent and can be realized by different algorithms. Tasks are taken from the aperiodic queue only when the periodic queue is empty. The activation of a new periodic instance causes any aperiodic tasks to be immediately preempted.



3.2 Immediate Execution

Contrary to the Background Execution, aperiodic tasks are served with the highest priority as soon as they come. This however, might cause deadline misses among the periodic tasks.



Aperiodic Servers are the solution to the problem. Normally we associate two parameters with a server:

- C_s: capacity
- T_s: server period

Roughly speaking, the idea is that the served tasks receive no more that C_s time units every T_s. How this is done depends on the specific server technology.

The server is scehduled as any periodic tasks. Priorities are manipulated in favour of the server. Tasks inside the server can be queued with an arbitrary discipline.

3.3 Polling Servers (PS)

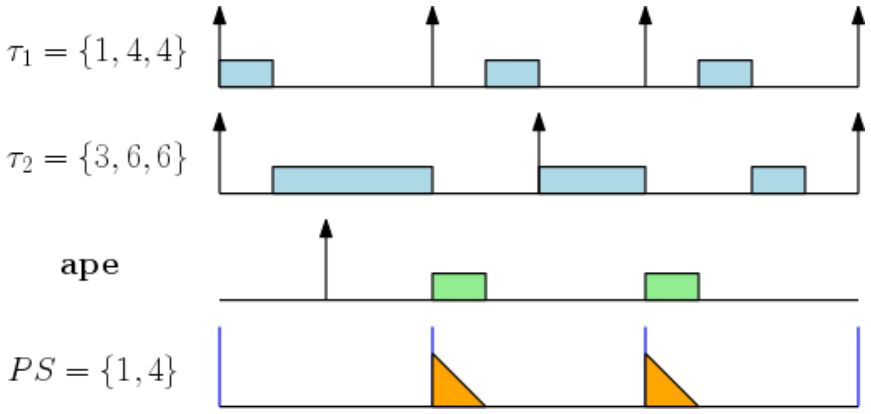
The average response time of aperiodic tasks can be improved with respect to background scheduling through the use of a **server**, that is, a periodic task whose purpose is to service aperiodic requests as soon as possible. Like any periodic task, a server is characterized by a **server period** T_s and a computation time C_s, called **server capacity**, or **server budget**. In general, the server is scheduled with the same algorithm used for the periodic tasks, and once active, it serves the aperiodic requests within the limit of its budget. The ordering of aperiodic requests does not depend on the scheduling algorithm used for periodic tasks, and it can be done by arrival time, computation time, deadline or any other parameter.

server

server period

server capacity

server budget



The **Polling Server (PS)** is an algorithm based on such an approach. At regular intervals equal to the period T_s, PS become active and serves the pending aperiodic requests withing the limit of its capacity C_s. If no aperiodic requests are pending, PS suspends itself until the beginning of its next period, and the budget originally allocated for aperiodic service is discharged and given periodi tasks.

Polling Server (PS)

Note that if an aperiodic request arrives just after the server has suspended, it must wait until beginning of the next period, when the server capacity is replenished at its full value.

The impact of the Polling Server on the other tasks is evaluated using the U_{lub}. In the case of a task set of periodic tasks scheduled via RM priority assignment and the presence of a polling server:

$$U_{lub}^{PS+RM} = U_s + n\left[\left(\frac{2}{U_s+1}\right)^{1/n} - 1\right]$$

In addition if the server runs at the highest priority we can compute its response time as follows:

$$R_a = \Delta_a + C_a + F_a(T_s - C_s)$$

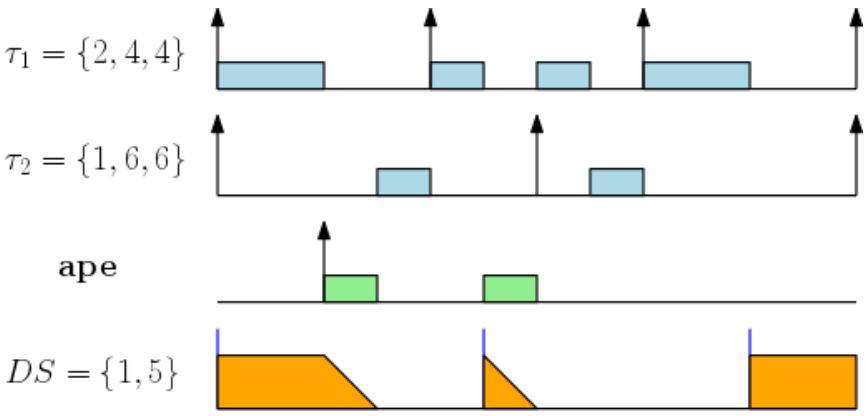
where:

- Δ_a is the initial delay
$$\Delta_a = \left\lceil \frac{r_a}{T_s} \right\rceil T_s - r_a$$
- F_a is the number of full service periods
$$F_a = \left\lceil \frac{C_a}{C_s} \right\rceil - 1$$
- δ_a is the final chunck
$$\delta_a = C_a - F_a C_s$$
- C_a is the computation time of the task
- C_s is the server capacity
- T_s is the server period

3.4 Deferrable Servers (DS)

The **Deferrable Server (DS)** algorithm is a service technique introduced by Lehoczky, Sha, and Strosnider to improve the average response time of aperiodic requests with respect to polling service. As the Polling Server, the DS algorithm creates a periodic task (usually having a high priority) for servicing aperiodic requests. However, unlike polling, DS preserves its capacity if no requests are pending upon the invocation of the server. The capacity is maintained until the end of the period, so that aperiodic requests can be services at the same server's priority at anytime, as long as the capacity has not been exhausted. At the beginning of any server period the capacity is replenished at its full value.

Deferrable Server (DS)



DS provides much better aperiodic responsiveness than polling, since it preserves the capacity until is needed. Shorter response times can be achieved by creating a Deferrable Server having the highest priority among the periodic tasks.

The impact of the Deferrable Server on the other tasks is evaluated using the U_{lub}. In the case of a task set of periodic tasks scheduled via RM priority assignment and the presence of a polling server:

$$U_{lub}^{DS+RM} = U_s + n\left[\left(\frac{U_s+2}{2U_s+1}\right)^{1/n} - 1\right]$$

3.5 Sporadic Servers (SS)

The **Sporadic Server (SS)** algorithm is another technique which allow the enhancement of the average response time of aperiodic tasks without degrading the utilization bound of the periodic task set.

Sporadic Server (SS)

The SS algorithm creates a high-priority task for servicing aperiodic requests and, like DS, preserves the server capacity at its high-priority level until an aperiodic request occurs. However, SS differs from DS in the way it replenishes its capacity. Whereas DS periodically replenish their capacity to full value at the beginning of each server period, SS replenishes its capacity only after it has been consumed by aperiodic task execution.

In order to simplify the description of the replenishment method used by SS, the following terms are defined:

- P_{exe} It denotes the priority level of the task that is currently executing
- P_s It denotes the priority level associated with SS
- **Active** SS is said to be active when $P_{exe} \geq P_s$
- **Idle** SS is said to be idle when $P_{exe} < P_s$
- **RT** It denotes the replenishment time at which the SS capacity will be replenished
- **RA** It denotes the replenishment amount that will be added to the capacity at time RT

Using this terminology, the capacity C_s consumed by aperiodic requests is replenished according to the following rules:

- The replenishment time RT is set as soon as SS becomes active and $C_s > 0$. Let t_a be such a time. The value of RT is set equal to T_a plus the server period

$$RT = t_a + T_s$$

- The replenishment amount RA to be done at time RT is computed when SS becomes idle or C_s has been exhausted. Let t_I be such ta time. The value of RA is set equal to the capacity consumed withing the interval $[t_a, t_I]$

3.6 Constant Bandwidth Servers (CBS)

In this section we present a novel service mechanism, called **Constant Bandwidth Server (CBS)**, which efficiently implements a bandwidth reservation strategy. The Constant Bandwidth Server guarantees that, if U_s is the fraction of processor time assigned to a server (i.e. its bandwidth), its contribution to the total utilization factor is no greater thatn U_s, even in the presence of overloads.

Constant Bandwidth Server (CBS)

The basic idea behind the CBS mechanism can be explained as follows: when a new job enters the system, it is assigned a suitable scheduling deadline (to keep its demand within the reserved bandwidth) and it is inserted in the EDF ready queue. If the job tries to execute more than expected, its deadline is postponed (i.e. its priority is decreased) to reduce the interference on the other tasks. Note that by postponing the deadline, the task remains eligible for execution. In this way, the CBS behaves as a work conserving algorithm, exploiting the available slack in an efficient (deadline-based) way, thus prividing better responsiveness with respect to non-work conserving algorithms and to other reservation approaches that schedule the extra portions of jobs in background.

If a subset of tasks is handled by a single server, all the tasks in that subset will share the same bandwidth, so there is no isolation among them. Novertheless, all the other tasks in the system are protected agains overruns occurring in the subset.

In order not to miss any hard deadline, the deadline assignment rules adopted by the server must be carefully designed.

Definition 20: Constant Bandwidth Server

A CBS is characterized by three main quantities:

- an ordered pair (Q_s, T_s) assigned by the user.
  Where Q_s is the maximum budget and T_s is the period of the server. The ratio
  $$U_s = \frac{Q_s}{T_s}$$
  is denotes as the server bandwidth.
- The current budget q_s (initialized to 0) managed by the server.
- The scheduling deadline d_s (initialized to 0) managed by the server.

Each served job J_k is assigned a dynamic deadline equal to the current server deadline. Whenever a served job executes, the server budget q_s is decreased by the same amount.

The CBS acts considering the following procedures:

1. When the server budget is exhausted (i.e. $q_s = 0$), the server budget is recharged at the maximum value Q_s and a new server deadline is generated as $d_s = d_s + T_s$. Note that there are no finite intervals of time in which the budget is equal to zero.
2. When a job J_k arrives and the server is active the request is enqueued in a queue of pending jobs according to a given (arbitrary) discipline.
3. When a job J_k arrives and the server is idle, if $q_s \geq (d_s - r_k)U_s$ the server generates a new deadline $d_s = r_k + T_s$ and q_s is recharged at the maximum value Q_s, otherwise the job is served with the last server deadline d_s using the current budget.
4. When a job finishes, the next pending job, if any, is served using the current budget and deadline. If there are no pending jobs, the server becomes idle.

Hence, the server behaviour can be described by the algorithm:

At arrival of job J_k at time r_k $\rightarrow$ Assign d_s
if $\exists$ pending aperiodic request **then**
  enqueue J_k
else
  if $(q_s \geq (d_s - r_k)\, U_s)$ **then**
    $q_s \leftarrow Q_s$
    $d_s \leftarrow r_k + T_s$
  else
    Continue to use the budget q_s with deadline d_s
  end if
end if

3.6.1 CBS Properties

The proposed CBS service mechanism presents some interesting properties that make it suitable for supporting applications with highly variable computation times. The most important one, the **temporal isolation property**, is formally expressed as follows:

temporal isolation property

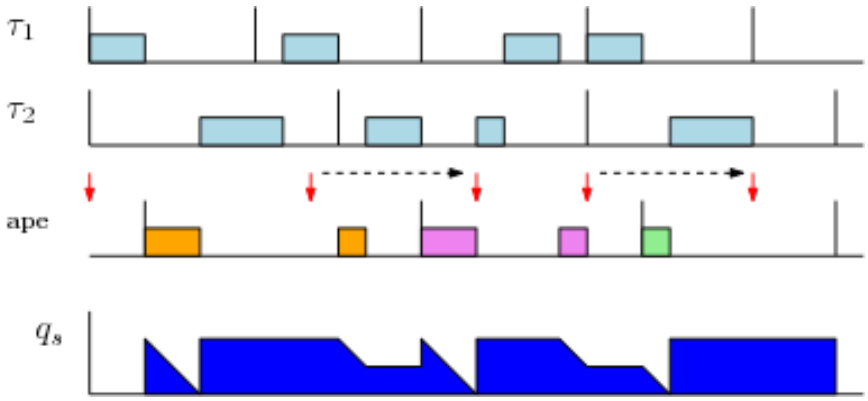


Figure 3.1: Example of two periodic task and CBS with $(Q_s, T_s) = (2, 6)$. The priority ordering is $\tau_1 > CBS > \tau_2$

Theorem 1

The CPU utilization of a CBS with parameters (Q_s, T_s) is

$$U_s = \frac{Q_s}{T_s}$$

independently from the computation times and the arrival pattern of the served jobs.

Lemma 1

Given a set of n periodic hard tasks with processor utilization U_p and a set of m CBSs with processor utilization

$$U_s = \sum_{i=1}^{m} U_{si}$$

the whole set is schedulable by EDF if and only if

$$U_p + U_s \leq 1$$

The temporal isolation property allows us to use a bandwidth reservation strategy to allocate a fraction of the CPU time to soft tasks whose computation time cannot be easily bounded. The most imporant consequence of this result is that soft tasks can be scheduled together with hard tasks without affecting the a priori guarantee, even in the case in which the execution time of the soft tasks are not known or the soft requests exceed the expected load.
Another general technique used in real-time systems for limiting the effects of overruns in tasks with variable computation times is the **resource reservation paradigm**. According to this method, each task is assigned a fraction of the processor bandwidth, just enough to satisfy its timing containts. The kernel, however, must prevent each task from consuming more that the requested amount to protect the other tasks in the systems (**temporal protection**). In this way, a task receiving a fraction U_i of the total processor bandwidth behaves as it were executing alone on a slower processor with a speed equal to U_i times the full speed. The advantage of this method is that each task can be guaranteed in isolation, independently of the behavior of the other tasks.

resource reservation paradigm

temporal protection

A simple and effective mechanism for implementing resource reservation in a real-time system is to reserve each task τ_i a specified amount of CPU time Q_i in every reservation period T_s.

Chapter 4

Resource Access Protocols

4.1 Introduction

A **resource** is any software structure that can be used by a process to advance its executino. resource
Tipically, a resource can be a data structure, a set of variables, a main mamory area, a file, or a set of registers of a peripheral device. A resource dedicated to a particular process is said to be *private*, whereas a resource that can be used by more tasks is called a *shared resource*. A shared resource protexted against concurrent accesses is called an *exclusive resource*.

To ensure consistency of the data structures in exclusive resources, any concurrent operating system should use appropriate resource access protocols to guarantee a mutual exclusion among competing tasks. A piece of code executed under mutual exclusion contraints is called a **critical section**. critical section

Any task that needs to enter a critical section must wait until no other task is holding the resource. A task waiting for exclusive resource is said to be *blocked* on that resource, othertwise it proceeds by entering the critical section and holds the resource. When a task leaves a critical section, the resource associated with the critical section becomes *free*, and it can be allocated to another waiting task, if any.

In this chapter, we describe the main problems that may arise in a uniprocessor system when concurrent tasks use shared resources in exclusive mode, and we present some resource access protocols designed to avoid such problems and bound the maximum blocking time of each task. We then show how such blocking times can be used in the schedulability analysis to extend the guaranteee tests derived for periodic task sets.

4.1.1 Atomicity

So far we have assumed that all tasks that run and compete for a processor are independent, which means that they do not interact with one another. But there are several occasions in which this assumption cannot be made (e.g. when two tasks need to share information, exchange variables,...). The other is when you have to compete for shared resources.

In this section we will introduce this problem and in particular we will introduce the notion of atomicity.

Definition 21: Atomic Instruction

An atomic instruction is an instruction whose execution cannot be interleaved with the execution of other instructions

In this sense atomic operations are always sequentialized since they cannot be interrupted. Under these conditions they are safe operations.
On the other hand, non atomic operations can be interrupted, and as such they are not *safe* operations.

Usually, it is preferrable to have, whenever possible, non atomic operations, because they allow you to exploit in full the possibility of scheduling the processor to activities having higher priority via preemption.

Example 11: Non atomic operations

Conside a simple operation like

$$x = x + 1$$

The variable is stored in a memory address that we call x. Hence, whenever the variable is incremented using this operation, we would have to:

- load the variable x into a register $R0$

```
LD R0, x
```

- increment the register

```
INC R0
```

- store back the value contained in the register into x

```
ST x, R0
```

If the same operation is executed inside an interrupt handler an inconsistency may arise.

Example 12: Interrupt on non-atomic operations

Let us consider that the increment operation is both applied in the normal code and in an interrupt handler code (routine executed in response to an interrupt).
In both cases, the operation is translated into the assembly language using three instructions (load, increment and store).
The program starts executing as follows:

1. The normal code starts executing: the value of x is loaded from memory to the register
2. At some point during this operation something triggers the execution of the interrupt
3. The interrupt handler creates a copy of all the registers
4. The interrupt handler load (once again) x from memory, increments the register and stores the result in memory (the value of x in memory has changed to $x + 1$)
5. Upon the interrupt handler has completed, the saved registers are restored. Hence, the old value of the register (i.e. x) is loaded back, its value is incremented to $x + 1$ and stored into memory at the address of x

The problem is that even though the code should have performed two increments (i.e. the final value should have been $x + 2$), it yields the incorrect result (i.e. $x + 1$).
From a logical point of view two increment operations should have taken place, but in effect one of them not successfully completed because while i was incrementing the variable, I was allowed to be interrupted and I was left with a state that was not up to date.

The nasty problem about this phenomenon is that it does not always happen in this way, because sometimes the function successfully completes before the interrupt is fired.
The example provided is the description of a condition called **critical race**, because you can have multiple execution of your code that interleave the operation in slightly different ways and you obtain different results. `critical race`

This is a nasty problem in computer science because it might not materialize for years.

This is so because you cannot make assumption about the speed of the hardware and on the exact moment when certain events take place (we do not know the order of execution of the hardware instructions).

The case studies proposed are a perfect example of a not atomic operation that should be atomic.

The same behaviour occurs not only in the case of interrupts but also on interleaving tasks: so you could have two tasks running in parallel, both of which are sharing the variable x.

We can give a few definitions that are important for the follow up of our discussion:

Definition 22: Shared Object

An object where the conflict may happen.

Definition 23: Critical section

A critical section is a sequence of operations that cannot be interleaved with other operations on the same resource

Definition 24: Mutual exclusion

Two critical sections cannot be active at the same time (they must be sequentialized). Either one or the other needs to stand by while the other executes

There are three ways to obstain mutual exclusion:

1. Implementing the critical section as an atomic operation.
   This that interrupt are disabled before the execution of the critical section and then are restored upon termination of the operation. The problem with this is that it is really tough because disabling the interrupts the I/O system of the machine is no longer allowed to work properly. (pressing an emergency button will have no effect and the execution of the code will continue),
   Moreover if the critical section is long, no interrupt can arrive during the critical section. In the case of a timer interrupt that arrives every 1ms, if a critical section lasts more than 1ms, a timer interrupt could be lost!

2. Disabling the preemption (system-wide).
   This strategy will have some problems, because all the tasks will suffer from this suspention of the preemption even though they do not use shared resources.

3. Selectively disabling the preemption (using semaphores and mutual exclusion).
   This strategy will disable preemption only for the task that operates on the shared resource.

Hence, we should try to disable preemption rather than disabling interrupts.

Still the big issue with selectively disabling preemption the priority mechanism is no longer enforced: during the critical section might force the process to execute low priority tasks because preemption is disabled. This phenomenon is known as **Priority Inversion**.

Priority Inversion

If Priority inversion is not correctly managed it may lead to a complete violation of all timing contraints.

4.1.2 Interacting Tasks

Until now, we have considered only independent tasks, which are characterized by the fact that a job never blocks or suspends and a task only blocks on job termination.
In the real world, jobs might block for various reasons:

- Tasks exchange data throught shared memory (mutual exclusion)
- A task might need to synchronize with other tasks while waiting for some data
- A job might need a hardware resource which is currently not available.

Example 13: Control Application

Let us consider a control application composed by three periodic tasks:

- τ_1 reads the data from the sensors and applies a filter. The results are stored in memory.
- τ_2 reads the filtered data and computes some control law (updating the state and the outputs); both the state and the outputs are stored in memory
- τ_3 reads the outputs and writes on an actuator

All of the three tasks access data in shared memory. This means that there are conflicts on accessing this data concurrently with the risk that the data structures become inconsistent.

4.1.3 Priority Inversion Phenomenon

The rest of this chapter presents the following resource access protocols:

- Non-Preemptive Protocol (NPP)
- Highest Locking Priority (HLP)
- Priority Inheritance Protocol (PIP)
- Priority Ceiling Protocol (PCP)

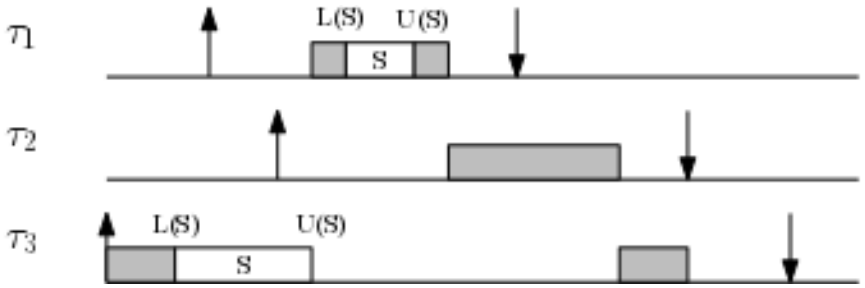
4.2 Non Preemptive Protocol (NPP)

A simple solution that avoids the unbounded priority inversion problem is to disallow preemption during the execution of any critical section. This method, also referred to as **Non-Preemptive Protocol (NPP)**, can be implemented by raising the priority of a task to the highest priority level whenever it enters a shared resource. In particular, as soon as a task τ_i enters a resource R_k, its dynamic priority is raised to the level:

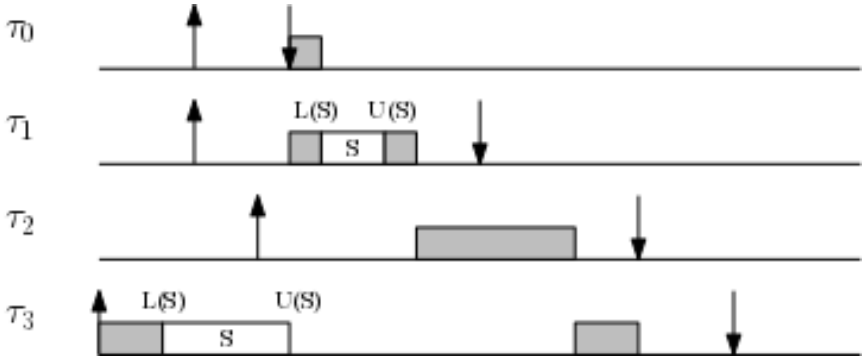
Non-Preemptive Protocol (NPP)

$$p_i(R_k) = \max_h \{p_h\}$$

The dynamic priority is then reset to the nominal value p_i when the task exits the critical section.



This method solves the priority inversion phenomenon, however, is only appriopriate when tasks use short critical sections because it creates unnecessary blocking. This actually might cause deadline misses of tasks that do not use the shared resource.



In the example, τ_0 misses its deadline (suffers a blocking time equal to 3) even though it does not use any resource!!
The solution is to raise τ_3 priority to the maximum between tasks accessing the shared resource (i.e. τ_1) priority.

4.2.1 Blocking Time and Response Time

NPP introduces a blocking time on all tasks bounded by the maximum length of a critical section used by lower priority tasks. Such blocking time affect the response times of the task as follows:

$$R_i = C_i + B_i + \sum_{j=1}^{i-1} \left\lceil \frac{R_i}{T_j} \right\rceil C_j$$

where:

- B_i is the blocking time from lower priority tasks
- $\sum_{h=1}^{i-1} \left\lceil \frac{R_i}{T_h} \right\rceil C_h$ is the interference from higher priority tasks

Example 14: Response time computation

Consider the following task set:

- $\tau_1 = (20, 30, 70)$ accessing resource for $\xi_{1,1} = 0$ units of time
- $\tau_2 = (20, 45, 80)$ accessing resource for $\xi_{2,1} = 1$ units of time
- $\tau_3 = (20, 130, 200)$ accessing resource for $\xi_{3,1} = 2$ units of time

It follows that the blocking times associated with each task is the maximum length of a critical section used by lower priority tasks:

- $B_1 = 2$ since τ_3 might block the task for 2 units of time when accessing its resource
- $B_2 = 2$ since τ_3 might block the task for 2 units of time when accessing its resource
- $B_3 = 0$ since there are no lower priority tasks that can block it

Thus, the response time of each task takes the form

$$R_i = C_i + B_i + \sum_{j=1}^{i-1} \left\lceil \frac{R_i}{T_j} \right\rceil C_j$$

- For τ_1

$$R_1^{(0)} = 20 + 2 = 22$$

- For τ_2

$$R_2^{(0)} = 20 + 2 = 22$$
$$R_2^{(1)} = 20 + 2 + \left\lceil \frac{22}{70} \right\rceil 20 = 42$$
$$R_2^{(2)} = 20 + 2 + \left\lceil \frac{42}{70} \right\rceil 20 = 42$$

- For τ_2

$$R_3^{(0)} = 35 + 0 = 35$$
$$R_3^{(1)} = 35 + 0 + \left\lceil \frac{35}{70} \right\rceil 20 + \left\lceil \frac{35}{80} \right\rceil 20 = 75$$
$$R_3^{(1)} = 35 + 0 + \left\lceil \frac{75}{70} \right\rceil 20 + \left\lceil \frac{75}{80} \right\rceil 20 = 95$$
$$R_3^{(1)} = 35 + 0 + \left\lceil \frac{95}{70} \right\rceil 20 + \left\lceil \frac{95}{80} \right\rceil 20 = 115$$
$$R_3^{(1)} = 35 + 0 + \left\lceil \frac{115}{70} \right\rceil 20 + \left\lceil \frac{115}{80} \right\rceil 20 = 115$$

4.3 Highest Locking Priority (HLP)

The **Highest Locking Priority (HLP)** protocol improves NPP by raising the priority of a task that enters a resource R_k to the highest priority among the tasks sharing that resource. In particular as soon as a task τ_i enters a resource R_k, its dynamic priority is raised to the level

Highest Locking Priority (HLP)

$$p_i(R_k) = \max_h \{p_i | \tau_h \text{ uses } R_k\} \tag{4.1}$$

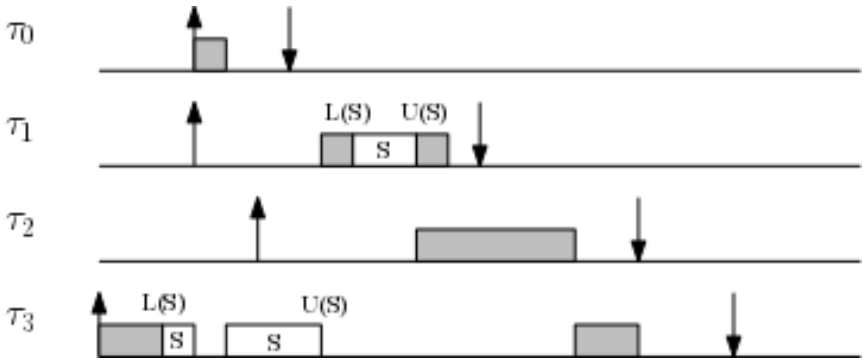
The dynamic priority is then reset to the nominal value p_i when the task exits the critical section. The online computation of the priority level in equation 4.1 can be simplified by assigning each resource R_k a **priority ceiling** $C(R_k)$ (computed offline) equal to the maximum priority of the tasks sharing R_k; that is:

priority ceiling

$$C(R_k) = \max_h \{p_i | \tau_h \text{ uses } R_k\}$$

Then, as soon as a task τ_i enters a resource R_k, its dynamic priority is raised to the ceiling of the resource. For this reason, this protocol is also referred to as **Immediate Priority Ceiling**.

Immediate Priority Ceiling



The main problem with HLP is that we must know in advance which task will access the resource: hence you need to have a complete control over the application, and how the application is written. Because knowing the code is the only way to know which resources is going to use. So this mechanism works very well if you have full control of what is running in your PC or embedded system.

4.4 Priority Inheritance Protocol (PIP)

The **Priority Inheritance Protocol (PIP)** proposed by Sha, Rajkumar and Lehoczky, avoids unbounded priority inversion by modifying the priority of those tasks that cause blocking. In

Priority Inheritance Protocol (PIP)

particular, when a task τ_i blocks one or more higher-priority tasks, it temporarily assumes (*inherits*) the highest priority of the blocked tasks. This prevents medium-priority tasks from preempting τ_i and prolonging the blocking duration experienced by the higher-priority tasks.

The Priority Inheritance Protocol can be defined as follow:

- Tasks are scheduled based on their active priorities. Tasks with the same priority are executed in a First Come First Served discipline.
- When task τ_i tries to enter a critical section and resource R_k is already held by a lower-priority task τ_j, then τ_i is blocked. τ_i is said to be blocked by the task τ_j that holds the resource. Otherwise, τ_i enters the critical section.
- When a task τ_i is blocked, it transmits its active priority to the task τ_j that holds the semaphore/mutex. Hence, τ_j resumes adn executes the rest of its critical section with a priority $p_j = p_i$. Task τ_j is said to inherit the priority of τ_i. In general, a task inherits the highest priority of the tasks it blocks. That is, at every instant,

$$p_j(R_k) = \max\{P_j, \max_h\{P_h | \tau_h \text{ is blocked on } R_k\}\} \tag{4.2}$$

- When τ_j exits a critical section, it unlocks the mutex/semaphore, and the highest-priority task blocked, if any, is awakened. Moreover, the avtive priority of τ_j is updated as follows: if no other tasks are blocked by τ_j, p_j is set to its nominal priority P_j; otherise it is set to the highest priority of the tasks blocked by τ_j, according to equation 4.2.
- Priority inheritance is transitive; that is, if a task τ_3 blocks a task τ_2, and τ_2 blocks a task τ_1, then τ_3 inherits the priority of τ_1 via τ_2

A high priority task can experience two kinds of blocking: **Direct blocking** and **Push-through blocking**.

Direct blocking

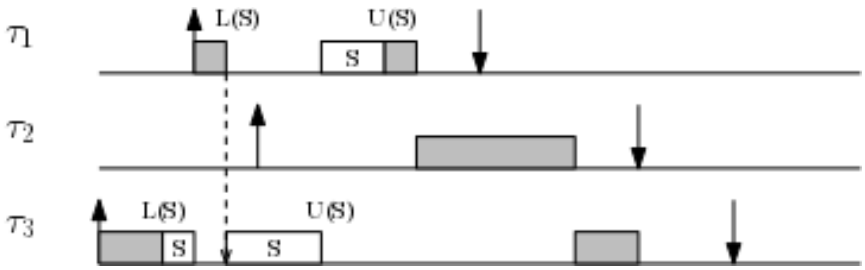
Push-through blocking

Definition 25: Direct blocking

It occurs when a higher-priority task tries to acquire a resource already held by a lower-priority task. Direct blocking is necessary to ensure the consistency of the shared resources

Definition 26: Push-through blocking

It occurs when a medium-priority task is blocked by a low-priority task that has inherited a higher priority from a task it directly blocks. Push-through blocking is necessary to avoid unbounder priority inversion



Although the Priority Inheritance Protocol bounds the priority inversion phenomenon, the blocking duration for a task can still be substantial because a chain of blocking can be formed. Another problem is that protocol does not prevent deadlocks (however, the latter problem can be solved by imposing a total ordering on the mutex accesses).

4.4.1 Blocking time and computation time

For the Priority inheritance protocol we only consider non nested critical sections. In fact, in presence of multiple inheritance, the computatino of the blocking time becomes very complex, whereas in non nested critical sections, multiple inheritance cannot happend and the computation of the blocking time becomes simpler.

The maximum blocking time can be computed based on two important properties which provide an upper bound on the number of times a task can block.

Theorem 2

If PI is used, a task block only once on each different critical section

Theorem 3

If PI is used, a task can be blocked by another lower priority task for at most the duration of one critical section

These two properties imply that a task can be blocker more than once, but only once per each resource and once by each task.

The Blocking time computation makes use of a **resource usage table**:

resource usage table

- A task per row, in decrasing order of priority
- A resource per column
- Cell (i, j) contains $\xi_{i,j}$, i.e. the lenght of the longest critical section of task τ_i on resource S_j, or 0 if the task does not use the resource.

The computation of the blocking time makes use of the resource usage table and follows the following procedure (taking into considerations the two PI properties):

- A task can be blocked only by lower priority tasks: then, for each task (row), we must consider only the rows below (tasks with lower priority)
- A task block only on resources directly used, or used by higher priority tasks (**indirect blocking**): for each task, only consider columns on which it can be blocked (used by itself or by higher priority tasks)

indirect blocking

Let us consider the following resource usage table:

	S_1	S_2	S_3
τ_1	2	0	0
τ_2	0	1	0
τ_3	0	0	2
τ_4	3	3	1
τ_5	1	2	1

Example 15

- B_1
  τ_1 can be blocked only on S_1. Therefore, we must consider only the first column, and take the maximu, which is 3.
  Therefore $B_1 = 3$.
- B_2
  τ_2 can be blocked on S_1 (indirect blocking) by lower priority tasks inheriting a higher priority and on S_2 by lower priority tasks.
  Consider all cases where two distinct lower priority tasks in $\{\tau_3, \tau_4, \tau_5\}$ access S_1 and S_2, sum the two contributions, and take the maximum:
  - τ_4 on S_1 and τ_5 on S_2: blocking time of 5
  - τ_4 on S_2 and τ_5 on S_1: blocking time of 4

  Hence, $B_2 = 5$
- B_3

τ_3 can be blocked on all 3 resources

- τ_4 on S_1 and τ_5 on S_2: blocking time of 5
- τ_4 on S_1 and τ_5 on S_3: blocking time of 4
- τ_4 on S_2 and τ_5 on S_1 or S_3: blocking time of 4
- τ_4 on S_3 and τ_5 on S_2: blocking time of 3
- τ_4 on S_3 and τ_5 on S_1: blocking time of 2

Hence, $B_3 = 5$

- B_4
  τ_4 can be blocked on all 3 resources, since it can be blocked only by τ_5
  - τ_5 on S_2: blocking time of 2
  - τ_5 on S_1 or S_3: blocking time of 1

  Hence, $B_4 = 2$
- τ_5 cannot be blocked by any other task (because it is the lower priority task).
  Hence, $B_5 = 0$

As usual the schedulability tests with a blocking time can be carried on using the three tests that we have seen:

- Processor utilization factor test.
  The system is schedulable if

$$\forall i \in [1, n] \qquad \sum_{k=1}^{i-1} \frac{C_k}{T_k} + \frac{C_i + B_i}{T_i} \leq i(2^{1/i} - 1)$$

- Response time analysis

$$R_i = C_i + B_i + \sum_{h=1}^{i-1} \left\lceil \frac{R_i}{T_h} \right\rceil C_h$$

- Processor demand analysis.
  In a task set $\mathcal{T}$ composed of independent and periodic tasks, τ_i is schedulable (for all possible phasing) if and only if

$$\exists t \in [0, D_i] \qquad W_i(0, t) = C_i + \sum_{h=1}^{i-1} \left\lceil \frac{t}{T_h} \right\rceil C_h \leq t - B_i$$

  As usual we can define

$$W_i(t) = C_i + \sum_{h=1}^{i-1} \left\lceil \frac{t}{T_h} \right\rceil C_h$$

$$L_i(t) = \frac{W_i(t)}{t}$$

$$L_i = \min_{t \in [0, D_i]} L_i(t) + \frac{B_i}{t}$$

  The task set is schedulable if $\forall i$, $L_i \leq 1$.
  Again, we can compute L_i by only considering the scheduling points.

4.5 Priority Ceiling Protocol (PCP)

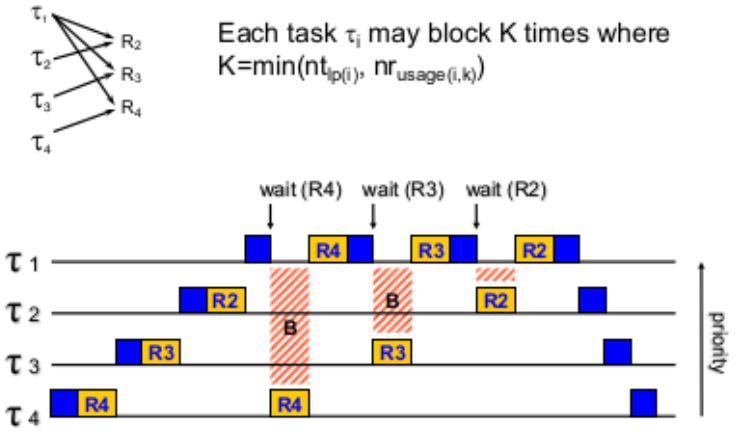
The **Priority Ceiling Protocol (PCP)** was introduced by Sha, Rajkumar, and Lehoczky to bound the priority inversion phenomenon and prevent the formation of deadlocks and chained blocking.

Priority Ceiling Protocol (PCP)

The basic idea of this method is to extend the Priority Inheritance Protocol with a rule granting a lock request on a free mutex. To avoid multiple blocking, this rule does not allow a task to enter a critical seciton if there are locked mutexes that could block it. This means that once a task enters its first critical section, it can never be blocked by lower-priority tasks until its completion.

In order to realize this idea, each mutex is assigned a priority ceiling equal to the highest priority of the tasks that can lock it. Then, a task τ_i is allowed to enter a critical section only if its priority is higher than all priority ceiling of the mutexes currently locked by tasks other than τ_i.

The problem with Priority Inheritance Protocol resides in the case where there are multiple blockings.



Example 16: Multiple Blocking

Consider a task set of four tasks $(\tau_1, \tau_2, \tau_3, \tau_4)$ sharing three resources R_2, R_3, R_4 in the following manner:

- τ_1 uses R_2, R_3, R_4
- τ_2 uses R_2
- τ_3 uses R_3
- τ_4 uses R_4

Now consider the following case:

1. The schedule starts executing τ_4
2. The task enters the critical section and locks resource R_4.
3. τ_4 is preempted, and τ_3 starts executing (higher priority task)
4. τ_3 enters a critical section and locks resource R_3.
5. τ_3 is preempted, and τ_2 starts executing (higher priority task)
6. τ_2 enters a critical section and locks resource R_2.
7. τ_2 is preempted, and τ_1 starts executing (higher priority task)
   At this point all three resources are locked by the three lower priority tasks
8. whenever τ_1 tries to access any of the three resources it is blocked and have to give back the execution to the task holding the resource.

In the case described above, the protocol perform three context switches between an high priority task and a lower priority task holding the resource introducing unnecessary blocking time to task τ_1 and performing too many context switches (which are expensive from a real-time perspective, to be precise 2 context switches for each blocking time).
In addition, the lower priority tasks preempted are active all the time, and therefore they

allocate on the stack their local variable. Within the stack we would have all the cumulative use of memory of each task at the same time.

Hence we can notice that Priority Inheritance Protocol brings several drawbacks:

- Tasks may block multiple times
- Worst case behaviour even worse that non-preemptable critical section
- Costly implementation except for very simple cases
- Does not even prevent deadlocks.

For all these reasons, priority ceiling was proposed.

Definition 27: Priority ceiling of a resource S

Maximum priority among all tasks that can psssibly access S

The mechanism behind the Priority Ceiling Protocol is that a process can only lock a resource if its dynamic priority is higher than the ceiling of any currently locked resource (excluding any that it has already locked itself). Thus if task τ blocks, the task holding the lock on the blocking resource inherits its priority.

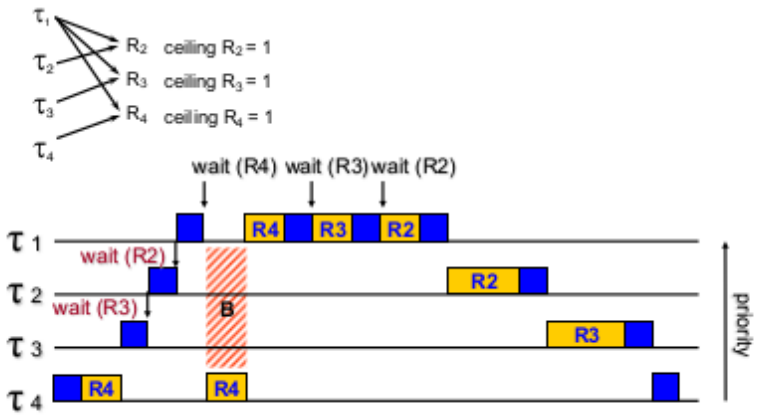
There are two forms of Priority Ceiling Protocol:

- Original Priority Ceiling Protocol (OPCP)
- Immediate Priority Ceiling Protocol (IPCP)

The nice priorities of both these protocols are:

- A high priority process can be blocked at most once during its execution by lower priority processes
- Deadlocks are prevented
- Transitive blocking is prevented, hence multiple inheritance cannot happen

4.5.1 Original Priority Ceiling Protocol (OPCP)



Example 17: Original Priority Ceiling Protocol

Consider a task set of four tasks $(\tau_1, \tau_2, \tau_3, \tau_4)$ sharing three resources R_2, R_3, R_4 in the following manner:

- τ_1 uses R_2, R_3, R_4

- τ_2 uses R_2
- τ_3 uses R_3
- τ_4 uses R_4

Given this resource utilization graph we can notice that all three resources have a priority ceiling of 1 since all three resources are used by task τ_1 with the highest priority among all the tasks.
Now consider the following case:

1. The schedule starts executing τ_4
2. The task enters the critical section and locks resource R_4.
3. τ_4 is preempted, and τ_3 starts executing (higher priority task)
4. τ_3 executes and tries to lock resource R_3, but the protocol prevents it from locking the resource. This is because their priority of τ_3 is lower than the Priority ceiling of resource R_4
5. τ_2 executes and tries to lock resource R_2, but the protocol prevents it from locking the resource. This is because their priority of τ_2 is lower than the Priority ceiling of resource R_4
6. τ_1 executes and tries to lock resource R_4, but the protocol prevents it from locking the resource. It then transmits its priority to τ_4
7. τ_4 resumes its utilization of R_4 and unlocks the resource
8. At this point τ_1 is the highest priority and can lock resource R_4, as well as the other resources.

We can see that the priority ceiling on resource R_4 prevented task τ_2 and τ_3 from locking resources R_2 and R_3.

This process allows to lock only once on one resource. Hence there are no excessive context switches and blocking time but still the stack occupation of the active tasks is the same.

4.5.2 Immediate Priority Ceiling Protocol (IPCP)

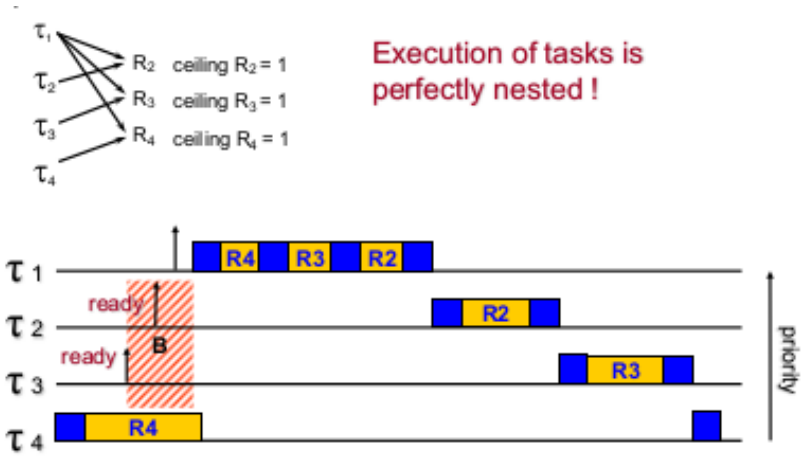
The Immediate priority ceiling considers the case where high and low priority tasks share a critical seciton. Still the ceiling is the highest priority among all tasks that can use the resource but this time not only a task is not allowed to occupy a resource if the priority is lower than the ceiling of any locked resource, but the task is not even allowed to start.

Example 18: Immediate Priority Ceiling Protocol

In the example proposed before:

1. The schedule starts executing τ_4
2. The task enters the critical section and locks resource R_4.
3. Task τ_3 (and τ_2, τ_1) tries to execute, but the protocol prevents it from doing so because its priority is lower than the priority ceiling of the locked resource
4. Hence τ_4 continues to utilize resource R_4 and upon releasing it, task τ_1 has the highest priority among all tasks and can utilize all three resources without context switches.

The immediate priority ceiling protocol prevents deadlocks, excessive context switches and blocking time and considerably lowers the stack occupation of the active tasks since they are reduced



(they are not even allowed to execute)

4.5.3 OPCP vs IPCP

The comparison between the two protocols is as follows:

- The worst case behaviour is identical from a scheduling point of view
- IPCP is easier to implement than the original OPCP as blocking relationships need not to be monitored
- IPCP leads to less context switches as blocking is prior to first execution
- IPCP requires more priority movements as this happens with all resource usages; OPCP only changes priority if an actual block has occurred
- IPCP allows substantial savings of stack occupation

4.5.4 Blocking time computation

Contrary to priority inheritance protocol, where the blocking time is the higher combination of resource utilization between the lower rows of the resource utilization table.
Whereas the Priority Ceiling Protocol only considers the maximum value of the submatrix of the lower rows of the resource utilization table.

Operating System Structure

Chapter 5

The Kernel

Recall the following elementary defintions:

Definition 28: Real-Time Operating Sytems (RTOS)

Operating System providing support to Real-Time applications

Definition 29: Rea-Time application

The correctness depends not only on the output values, but also on the time when such values are produced

Definition 30: Operating System

An Operating System is:

- Set of computer programs
- Interface between applications and hardware
- a way to control the execution of application programs
- a way to manage the hardware and software resources

In particular we can interpret an operating system as a:

- a Service Provider: an API plus some services behind. This API needs to be suitable for real-time application.
- a Resource Manager: implements schedulers, policies to share resource, aperiodic servers, ...

The services provided by the Operating System are executed in **Kernel Space**: whenever you execute a program on the operating systems the processor switches in a particular mode called **supervisor mode** (in this mode the processor can do anything allowed by the machine including interrupts, manage memory, ...). In this kernel space the operating system is able to:

`Kernel Space`

`supervisor mode`

- Process Synchronization, Inter-Process Communication
- Schedule processes/threads
- Input and Output
- Allocate Virtual Memory

All of these things are exported and accessible by means of an API.

5.1 Introduction

The core of the machine is called the **Kernel**

`Kernel`

Definition 31: Kernel

core part of the OS, allowing multiple tasks to run on the same CPU

The core feature of the kernel is that it allows a task set $\mathcal{T}$ composed by N tasks to run in parallel in a M CPUs ($M < N$). From the application/program point of view there is no difference between having a task executed in an intermediate way or having tasks executed in a dedicated processor: so what the operating system does is to ensure a proper temporal multiplexing between the tasks.

The task scheduling service relies on two core components:

- **Scheduler**: decides which task to execute — Scheduler
- **Dispatcher**: component that implements the context switch between the tasks. Often this component relies on some hardware extension that facilitates this type of service. — Dispatcher

The kernel also provides a mechanism for allowing tasks to communicate and synchronize between each other. On this regard there are two possible paradigms:

- Shared memory (threads).
  Shared Memory utilizes mutexes, semaphores and condition variables, which the kernel provides. From a real time point of view this service has to come along some real-time resource sharing protocols.
- Message passing (processes).
  Contrary to Shared memory, Message passing is based on different interaction models such as pipeline, client-server, ...
  The kernel must once again provide some IPC mechanism: pipes, message queues, mailboxes, remote procedure calls (RPC), ...
  On top of this some real-time protocols can still be used

In practice, an adequate scheduling of system resources removes the need for over-enginnering the system, and is necessary for providing a predictable QoS. For this reason, an adequate scheduling of system resources considers two important aspects: the algorithm and the implementation.
For instance, we have seen efficient algorithm for scheduling tasks and resources, however all applications are treated by a **Timer**. This introduces a set of questions and problems to consider in the implementation of the scheduling algorithm: — Timer

- Is the timer reliable?
- Is the scheuduler able to select a high-priority task as soon as it is ready?
- And the dispatcher?

Example 19: Periodic Task

When considering a periodic task, it expects to be executed at time $r = r_0 + jT$, but sometimes it is delayed to $r = r_0 + jT + \delta$, where the quantity δ becomes an offset or a drift term that makes the timing not reliable.
This delay may cause deadline misses.

5.2 Kernel Latency

Definition 32: Kernel Latency

Delay δ with which a kernel implement its decisions

When a primitive is called, the primitive takes some time to execute and during this time there is not possibility preempt the kernel. Therefore, in practice, the operating system is temporarily sospending the scheduling mechanism. This is situation really similar to the resource access protocols,

in fact, we can think of the kernel as a shared resource that is not preemptable and therefore the kernel latency can be modelled as a blocking time.
Hence the schedulability analysis tools introduced are modified as follows:

- **Processor Utilization factor test**

$$\forall i \in [1,n] \qquad \sum_{k=1}^{i-1} \frac{C_k}{T_k} + \frac{C_i + \delta}{T_i} \leq U_{lub}$$

- **Response Time Analysis**

$$R_i = C_i + \delta + \sum_{h=1}^{i-1} \left\lceil \frac{R_i}{T_h} \right\rceil C_h$$

- **Processor Demand analysis**

$$\exists 0 \leq t \leq D_i \qquad W_i(0,t) = C_i \sum_{h=1}^{i-1} \left\lceil \frac{t}{T_h} \right\rceil C_h \leq t - \delta$$

The scheduler is called whenever an internal (e.g. IPC, signal, ...) or external (e.g. interrupt) events is triggered.
When one of these events take place, there exists some time between the triggering of the event and the dispatch which can be decomposed in four components:

- Event generation
- Event delivery (interrupts may be disabled)
- Scheduler activation (non preemptable section)
- Scheduling time

There are quite a few elements that may generate the delay.

All these elements of uncertainty need to be understood, the reason and how the system manages introduces this delay and most importantly how to manage this delay in a real-time kernel.
Remember that we are interested in the worst case possibility, so all the components that compose the kernel latency should be designed so that in the worst-case the maximum is minimized.

5.3 System Architecture

The system architecture is composed by a system bus that is interconnecting the following components:

- One or more CPUs
- Memory (RAM)
- I/O Devices
  - Secondary memory (disks, ...)
  - Network cards
  - Graphic cards
  - Keyboard, mouse, ...

5.3.1 The CPU

The model of the CPU is composed of the following registers

- General-purpose registers that can be accessed by all the programs. These registers can be either data registers or address registers
- Program Counter (PC) aka Intruction pointer
- Stack Pointer (SP) register
- Flags register (aka Program Status Word)
- Some special registers, which control how the CPU works, must be "protected"

Regual user programs should not be allowed to influence the CPU mode of operation, perform I/O operations and reconfigure virtual memory. For this reason there is a need for privileged mode of execution

- Regual registers vs special registers
- Regual intructions vs privileged instructions

User programs: low privilge level (**User Level**) User Level
The OS kernel runs in supervisor mode.

Example 20: Intel x86

Real CPUs are more complex. They have few General Purpose registers: EAX, EBX, ECX, EDX (accumulator registers containing an 8 bit part and a 16 bit part), EBP, ESI, EDI

- EAX: Main accumulator
- EBX: sometimes used as base for arrays
- ECX: sometimes used as counter
- EBP: stack base pointer (for subroutines calls)
- ESI: source index
- EDI: destination index

They also have segmented memory architecture: segment registers CS (code segment), DS (data segment), SS (stack segment), GS, FS.
Finally they have various mode of operation: RM, PM, VM86, x86-64,..., mainly due to backward compatibility.

The Kernel is part of the OS which manages the hardware.
Runs with the CPU in Supervisor Mode (high privilege level):

- Privilege level known as Kernel Level (KL), execution in Kernel Space
- Regual programs run in User Space

Mechanismns for increasing the privilege level (from US to KS) in a controlled way

- Interrupts (+ traps/hw exeptions).
- Instructions causing a hardware exception

Switch the CPU from User Level to Supervisor mode by entering the kernel. This can be used to implement system calls. A partical Context Switch is performed: flags and PC are pushed to the stack, if the processor is executing at User Level, switch to Kernel Level, and eventually switch to a kernel stack finally execution jumps to a handler in the kernel (save the user registers for restoring them later). Once finished return to low privilege level (execution returns to User Space) through a "return from interrupt" Assembly instruction (`IRET` on x86): Pop flags and PC from stack and eventually switch back to user stack. Return path from system calls and hardware interrupts to handlers.

To understand interrupts, copnsider simplified CPU execution first

The CPU interatively:

- Fetch an instruction (address given by PC)
- Increase the PC
- Execute the intruction (might update the PC on jump...)

A More realistic execution model

Interrupt cannot fire during the execurtion of an intruction. Hardware exception: cause by the execution of an intruction:

- `trap`,`syscall`,`sc`,...
- I/O instructions at low privilege level, Page faults, ...

The interrupt tabole holds the addresses of the handlers:

- Interrupt n fires: after eventually switching to KS and pushing flags and PC on the stack
- Read the address contained in the n^{th} entry of the interrupt table, and jump to it!

Interrupt tables are implemented in hardware or in software:

- x86, interrupt description table composed by interrupt gates. The CPU automatically jumpts to the n^{th} interrupt gate
- Other CPUs jump to a fixed address: a software demultiplexer reads the interrupt table

Software Interrupt - System Call

1. Task τ_1, executes and invokes a system call
2. Execution passes from US to KS (change stack, push PC and flags, increase privilege level)
3. The invoked syscall executes. Maybe, it is blocking
4. τ_1 blocks and the system returns to US, and τ_2 is scheduled

Hardware Interrupt

1. Task τ_2 is executing, a hardware interrupt fires
2. Execution passes from US to KS (change stack, push PC and flags, increase privilege level)
3. the proper Interrupt Service Routine executes
4. The ISR can unblock τ_1. When execution returns to US, τ_1 is scheduled

The executino flow enters the kernel for two reasons

- Reacting to events coming from up (syscalls)
- Reacting to an event coming from below (an hardware interrupt from a device)

The kernel executes in the context of the interrupted task.
A system call can block the invoking task, or can unblock a different task
An ISR can unblock a task
If a task is blocked/unblocked, when returning to user space a context switch can happen.
The scheduler is invoked when return from KS to US

Example 21: I/O operation

Consider a generic Input or Output to an external device such as a PCI card.
This operation is performed by the kernel and user programs must uyse a syscall
The operation is performed in 3 phases:

1. Setup: prepare the device for the I/O operation
2. Wait: wait for the end of the operation
3. Cleanup: complete the operation

This can be done using polling, PIO, DMA, ...

5.3.2 Polling

User programs invoke the kernel; execution in kernel space until the operation is terminated.
The kernel cyclically reads (polls) an interafce status register to check if the operation is terminated
Busy-waiting in kernel space!

- No user task can execute while waiting for the I/O operation ...
- The operation must be very short
- I/O operation == blocking time

1. The user program raises a software input
2. Setup phase - in kernel: in case of input operation, nothing is done; in case of output operation, write a value to a card register
3. Wait - in kernel: cycle until a bit of the card status register becomes 1
4. Cleanup - in kernel: in case of input, read a value from a card register; in case of output, nothing is done. Eventually return to phase 1
5. IRET

5.3.3 Programmed I/O

User programs invoke tyhe kernel; execution returns to user space while waiting for the device: the task that invoked the syscall blocks
Anb interrupt will notify the kernel when the "wait" phase is terminated:

- The interrupt handler will take care of performing the I/O operation
- Many frequent short interruption of unrelated user-space tasks

1. The user program raises a software input
2. Setup phase - in kernel: instruct the device to raise an input when it is ready for I/O
3. Wait - return to user space: block the invoking task, and schedule a new one (IRET)
4. Cleanup - in kernel: the interrupt fires $\rightarrow$ enter kernel, and perform the I/O operation
5. Return to phase 2, or unblock the task if the operation is terminated (IRET)

5.3.4 DMA

User programs invoke the kernel; execution returns to user space while waiting for the device. The task that invoked the syscall blocks!

I/O operations are not performed by the kernel on interrupt, Performed by a dedicated HW device,
An interrupt is raised when the whole I/O operation is terminated

1. The user program raises a software input
2. Setup phase - in kernel: instruct the DMA (or the Bus Mastering Device) to perform the I/O
3. Wait - return to user space: block the invoking task, and schedule a new one (IRET)
4. Cleanup - in kernel: the interrupt fires → the operation is terminated. Stop device and DMA
5. Unblock the task and invoke the scheduler (IRET)

Chapter 6

Timer and Clock Latency

The introduction of a Latency is a violation of our expectation of the task: the task expects to be scheduled at a certain time but in effect it is scheduled later.

Definition 33: Latency

Measure of the diffence between the theoretical and actual schedule.

Example 22

A task τ expects to be scheduled at time t, but is actually scheduled at time t'. The resulting latency L takes the form:

$$L = t' - t$$

The latency L can be modelled as a blocking time and as such affects the guarantee test and schedulability analysis. Similar to what done for shared resources.
Blocking time due to latency, not to priority inversion.

The problem is: are we able to compute an upper bound for this latency? Otherwise we are not able to apply schedulability tests. Hence, the latency must be bounded:

$$\exists L^{max} : L < L^{max}$$

Otherwise the system is not well-behaved from a real-time perspective.

Formally: consider task τ_i is a stream of jobs $J_{i,j}$ arriving at time $r_{i,j}$.
A job $J_{i,j}$ is schedulable at time $t' > r_{i,j}$. The difference $t' - r_{i,j}$ is composed of three pieces:

1. $J_{i,j}$'s arrival is signalled at time $r_{i,j} + L^1$
2. Such event is served at time $r_{i,j} + L^1 + L^2$
3. $J_{i,j}$ is actually scheduled at $r_{i,j} + L^1 + L^2 + L^3$

where:

- L^1 is due to the delayed interrupt generation.
  Remember that the hardware interrupts are generated by devices. The device should generate and interrupt at time t but due to some internal issues, this interrupt generation takes place at time $t' = t + L^{int}$ where L^{int} is the Interrupt Generation Latency.
  The nature of this latency is due to some hardware issues (there is some time for the hardware to switch the state of some variables and thereby make the processor realize that an interrupt is fired). It is generally small when compared to L^{np}, but there is a case in which this interrupt generation is extremely important: the **Timer Resolution Latency** L^{timer}/
- L^2 is the **Non-preemptable Section Latency** (L^{np}).
  The non-preemptable section latency is the delay between time when an event is generated and when the kernel handles it. Such delay can be further broken down in several pieces: interrupt disabling, bottom halves dealying,...

Timer Resolution Latency

Non-preemptable Section Latency

- L^3 is the **Scheduler Latency** or **Scheduler Interference**. Which is the interference from higher priority tasks and its already accounted by the guarantee tests. Hence it will not be considered.

Scheduler Latency

Scheduler Interference

6.1 Timer Resolution Latency

The latency introduced by the timer L^{timer} can often be much larger that the non-preemptable section latency L^{np} (normally is neglectable).
Where does it come from? Timers inside the kernel are activate by a hardware device that produces periodic interrupts. Can we do anything about it?

6.1.1 Timer

A timer device produces periodic interrupts known as **Tick**s.

Tick

Example 23

Periodic tasks relies on the timer device (`setitimer()`, Posix timers, `clock_nanosleep()`,...)
τ_i with period T_i

Hence, the tick generation is a basic service where at each tick the kernel checks if some of the task needs to be waken up. Activations are triggered by the periodic interrupt:

- Periodic tick interrupt, with period T^{tick}
- Every T^{tick}, the kernel checks if the task must be woken up
- If T_i is not multiple of T^{tick}, τ_i experiences a timer resolution latency

If the tick is an integer submultiple of the period there is no issue with the schedulability of the tasks, but if the tick is 5 ms and the period of the task is set at 12ms there is an issue. This is the reason why traditional timers are not suitable for real-time applications. In a traditional operating systems contains a programmable timer device in periodic mode and every tick an interrupt is fired the system enters in kernel mode. The kernel executes and can:

- Wake up tasks
- Adjust tasks priorities
- Run the scheduler, when returning to user space (possible preemption)

There is a clear trade-off between responsiveness (low latency) and throughput (low overhead).

- Large T^{tick}: large timer resolution latency
- Small T^{tick}: high number of interrupts, more switches between US and KS, tasks are interrupted more often and therefore the system will have a resulting large overhead

For non real-time systems, it is possible to find a reasonable trade-off but it still depends on the workload.

Example 24: Linux Kernel

- Linux 2.4: 10 ms (100Hz)
- Linux 2.6: 100 Hz, 250 Hz or 1000Hz
- Other systems: $T^{tick} = 1/1024$

6.1.2 Bounded Timer Resolution Latency

The timer resoultion latency is experienced by all tasks that want to sleep for a specified time T.
τ_i must wake up at time $r_{i,j} = jT_i$, but is woken up at time $t' = \left\lceil \frac{r_{i,j}}{T^{tick}} \right\rceil T^{tick}$.

The Timer Resolution Latency is bounded:

- $t = r_{i,j}$
- $t' = \left\lceil \frac{r_{i,j}}{T^{tick}} \right\rceil T^{tick}$

$$\begin{aligned} L^{timer} &= t' - r_{i,j} \\ &= \left\lceil \frac{r_{i,j}}{T^{tick}} \right\rceil T^{tick} - r_{i,j} \\ &= \left(\left\lceil \frac{r_{i,j}}{T^{tick}} \right\rceil - \frac{r_{i,j}}{T^{tick}} \right) T^{tick} \leq T^{tick} \end{aligned}$$

Tipically, tasks are activated every few milliseconds and it is desirable to have a tick period around 2-5 ms. However, reducing T^{tick} below 1ms is generally not acceptable, so periodic tasks can expect a blocking time due to L^{timer} up to 1ms. Hence it is desirable to have periods which are integer multiples of the tick period and this will put limitations on the minimum duration of a period.

Additional problems:

- Tasks' periods are rounded to multiples of T^{tick}
- Limit on the minimum task period: $\forall i, T_i \geq T^{tick}$
- A lot of useless timer interrupts might be generated

Hence, we need a different technology to use instead of the timer device.

6.1.3 Clocks

There is also another important problem: there is some misconception behind Timers and Clocks

Definition 34: Timer

A timer generates an event at a specified time t

Definition 35: Clock

A clock keeps track of the current system time

In other words, a clock is needed to keep track of the evolution of time, whereas a timer is used to wake up tasks. They are both importants for real time applications: clocks are needed to take precise measurements on the system, timers are needed to generate activation sequences.

Definition 36: Timer Resolution

minimum interval at which a periodic timer can fire. If periodic ticks are used, the timer resolution is T^{tick}

Definition 37: Clock Resolution

minimum difference between two different times returned by the clock.

What's the expected clock resolution?

- Traditional OSs use a "tick counter".
  OS have a very fast clock and the OS returns the number of ticks (jiffies in Linux) from the system boot. For this reason the clock resolution in these systems is: T^{tick}
- Modern PCs have higher resolution time sources...
  On x86, TSC (TimeStamp Counter) inside the processor is a counter that keeps track of the number of clock cycles have been generated since you have turned on the machine.
  This clocks are called High-resolution clock: use the TSC to compute the time since the last timer tick...and have a precision of around 1 ns.

- Even using a "traditional" periodic timer tick, it is easy to provide high-resolution clocks: time can be easily read with a high accuracy.
- On the other hand, timer resolution is limited by the system tick T^{tick}. It is impossible to generate events at arbitrary instants in time, without latencies.

6.2 Timer Devices

The solution to the timer latency is to use a **Timer Device**: instead of using a timer that simply generates interrupts at a specific instant of time, there are programmable timers that can work high resolution (microseconds resolution). Timer Devices (e.g. PIT - i8254) generally work in 2 modes: periodic and one-shot. Timer Device

Generally speaking the concept is really simple: this device has a counter and an internal oscillator, every time an oscillation has completed it decrements a counter and only when the counter reaches zero an interrupt is generated. In summary it is the device itself that takes care of decrementing the counter and when to generate the interrupt.

As said before Timer Devices operate in two modes:

- If the device is programmed in periodic mode, the counter register is automatically reset to the programmed value
- If the device is programmed in one-shot mode, the kernel has to explicitly reprogram the device (setting the counter register to a new value)

The periodic mode is easier to use! This is why most kernels use it.
When using one-shot mode, the timer interrupt handler must:

1. Acknowledge the interrupt handlerm, as usual
2. Check if a timer expired, and do its usual stuff...
3. Compute when the next timer must fire
4. Reprogram the timer device to generate an interrupt at the correct time

Steps 3 and 4 are particularly critical and difficult: when the kernel reprograms the timer device (step 4), it must know the current time, but the last known time is the time when the interrupt fired (before step 1).

- A timer interrupt fires at time t_1
- The interrupt handler starts (enter KS) at time t_1'
- Before returning to US, the timer must be reprogrammed, at time t_1''
- Next interrupt must fire at time t_2; the counter register is loaded with $t_2 - t_1$
- Next interrupt will fire at $t_2 + (t_1'' - t_1)$

The error described previously accumulates with the risk of drift between real time and system time.
A free run counter (not stopped at t_1) is needed.
The counter is synchronised with the timer device and the value of the counter at time t_1 is known.
This permits to know the time t_1''. The new counter register value can be computed correctly.
On a PC, the second PIT counter, or the TSC, or the APIC timer can be used as a free run counter.

Serious real-time kernels use high-resolution timers (use hardware time in one-shot mode) which for instance is already implemented in RT-Mach, RTLinux, RTAI and others.
General purpose kernels are more concerned about stability and overhead.

Compatibility with "traditional" kernels:

- The tick event can be emulated through high-resolution timers
- Timer device programmed to generate interrupts both: when needed to serve a timer and at tick boundaries but the "tick" concept is now useless (e.g. Tickless or NO_HZ system which are good for saving power)

Chapter 7

The Non Preemptable Section Latency

Definition 38: Non-Preemptable Section Latency

Delay between time when an event is generated and when the kernel handles it.

- Due to non-preemptable sections in the kernel, which delay the response to hardware interrupts
- Composed by various parts: interrupt disabling, bottom halves delaying,...
- Depends on how the kernel handles the various events

The non-preemptable section latency L^{np} is given by the sum of different components:

1. Interrupt disabling
2. Delayed interrupt service
3. Delayed scheduler invocation

The first two are mechanisms used by the kernel to guarantee the consistency of internal structures. The third mechanism is sometimes used to reduce the number of preemptions and increase the system throughput.

7.1 Interrupt disabling

Before checking ifg an interrupt is fired, the CPU checks if interrupts are enabled. Every CPU has some protected intructions (`STI/CLI` on x86) for enabling/disabling interrupts.
In modern system, only the kernel (or code running in KS) can enable/disable interrupts.
Interrupts disabled for a time $T^{cli} \rightarrow L^{np} \geq T^{cli}$
Interrupt disabling is used to enforce mutual exclusion between sections of the kernel and ISRs.

7.2 Dealyed Interrupt Service

When the interrupt fires, the ISR is ran, but the kernel can delay the interrupt service some more...

- ISRs are generally small, and do only few things
- An ISR can set some kind of software flag, to notify that the interrupt fired
- Later, the kernel can check such flag and run a larger (and more complex) interrupt handler

Hard IRQ handlers (ISRs) vs "Soft IRQ handlers".
The advantages of "soft IRQ handlers" are:

- ISRs generally run with interrupts disabled
- Soft IRQ handlers can re-enable hardware interrupts
- Enabling/Disabling soft handlers is simpler/cheaper

Disadvantages:

- Increase NP latency: $L^{np} >> T^{cli}$
- Soft IRQ handlers are often non-preemptable increasing the latency for other tasks too...

7.3 Delayed scheduler invocation

Scheduler invoked when returing from KS to US. Sometimes, return to US after a lot of activities

- Try to reduce the numbert of KS - US switches
- Reduce the number of context switches
- Throughput vs low latency

ISR executed at the correct time, soft IRQ handler ran immediately, but scheduler invoked too late

7.4 Summary

L^{np} depends on some different factors.
In general, no hardware reasons, it almost entirely depends on the kernel structure: non-preemptable section latency is generally the result of the strategy used by the kernel for ensuring mutual exclusion on its internal data structures.

To analyze/reduce L^{np}, we need to understand such strategies.
Different kernels, based on different structures, work in different ways.
Some activities causing L^{np} are: interrupt handling (device drivers) and management of the parallelism

Additional information and proofs

Appendix A

U_{lub} for RM for N tasks

Given a task set $\mathcal{T}$ of N tasks scheduled using the RM priority assignment, the conditions that allow to compute the least upper bound of the processor utilization factor are:

$$\begin{cases} T_1 < T_n < 2T_1 \\ C_1 = T_2 - T_1 \\ C_2 = T_3 - T_2 \\ \dots \\ C_{n-1} = T_n - T_{n-1} \\ C_n = T_1 - \sum_{i=1}^{n-1} C_i = 2T_1 - T_n \end{cases}$$

Thus the processor utilization factor becomes

$$T = \frac{T_2 - T_1}{T_1} + \frac{T_3 - T_2}{T_2} + \dots + \frac{T_n - T_{n-1}}{T_{n-1}} + \frac{2T_1 - T_n}{T_n}$$

Defining

$$R_i = \frac{T_{i+1}}{T_i}$$

and noting that

$$\prod_{i=1}^{n-1} R_i = \frac{T_n}{T_1}$$

the utilization factor may be written as

$$U = \sum_{i=1}^{n-1} R_i + \frac{2}{\prod_{i=1}^{n-1} R_i} - n$$

To minimize U over R_i, $i = 1, \dots, n-1$, we have

$$\frac{\partial U}{\partial R_k} = 1 - \frac{2}{R_i \prod_{i=1}^{n-1} R_i}$$

Thus defining $P = \prod_{i=1}^{n-1} R_i$, U is minimum when:

$$\begin{cases} R_1 P = 2 \\ R_2 P = 2 \\ \dots \\ R_{n-1} P = 2 \end{cases}$$

that is, when all R_i have the same value

$$R_1 = R_2 = \cdots = R_{n-1} = 2^{1/n}$$

Substituting this value in U we obtain

$$\begin{aligned} U_{lub} &= (n-1)2^{1/n} + \frac{2}{2^{(1-1/n)}} - n \\ &= n2^{1/n} - 2^{1/n} + 2^{1/n} - n \\ &= n(2^{1/n} - 1) \end{aligned}$$

Appendix B

U_{lub} for RM + Polling Server

We first consider the problem of guaranteeing a set of hard periodic tasks in the presence of soft aperiodic tasks handled by a Polling Server. Then we show how to derive a schedulability test for hard aperiodic requests.

The schedulability of periodic tasks can be guaranteed by evaluating the interference introduced by the Polling Server on periodic execution. In the worst case, such an interference is the same as the one introduced by an equivalent periodic task having a periodi equal to T_s and a computation time equal to C_s. In fact, independently of the number of aperiodic tasks handled by the server, a maximum time equal to C_s is dedicated to aperiodic requests at each server period. As a consequence, the processor utilization factor of the Polling Server is

$$U_s = \frac{C_s}{T_s}$$

and hence the schedulability of a periodic set with n tasks and utilization U_p can be guaranteed if

$$U_p + U_s \leq U_{lub}(n+1)$$

If periodic tasks (including the server) are scheduled by RM, the schedulability test becomes

$$\sum_{i=1}^{n} \left(\frac{C_i}{T_i} \right) + \frac{C_s}{T_s} \leq (2^{1/(n+1)} - 1)(n+1)$$

Note that more Polling Servers can be created and execute concurrently on different aperiodic task sets.
In general, in the presence of m servers, a set of n periodic tasks is schedulable by RM if

$$U_p + \sum_{j=1}^{m} U_s j \leq U_{lub}(n+m)$$

A more precise schedulability test can be derived by assuming that PS is the highest-prioooority task in the system. To simplify the computation, the worst-case relations among the tasks are first determined, and then the lower bound is computed against the worst-case model.

Consider a set of n periodic tasks $(\tau_1, \ldots, \tau_n)$ ordered by increasing periods, and a PS server with highest priority. The worst-case scenario for a set of periodic tasks that fully utilize the processor is characterized by the following parameters:

$$\begin{cases} C_s = T_1 - T_s \\ C_1 = T_2 - T_1 \\ C_2 = T_3 - T_2 \\ \ldots \\ C_{n-1} = T_n - T_{n-1} \\ C_n = T_s - C_s - \sum_{i=1}^{n-1} C_i = 2T_s - T_n \end{cases}$$

The resulting utilization is then

$$\begin{aligned} U &= \frac{C_s}{T_s} + \frac{C_1}{T_1} + \cdots + \frac{C_n}{T_n} \\ &= U_s + \frac{T_2 - T_1}{T_1} + \cdots + \frac{T_n - T_{n-1}}{T_{n-1}} + \frac{2T_s - T_n}{T_n} \\ &= U_s + \frac{T_2}{T_1} + \cdots + \frac{T_n}{T_{n-1}} + \left(\frac{2T_s}{T_1}\right)\frac{T_1}{T_n} - n \end{aligned}$$

Defining:

$$\begin{cases} R_s = \dfrac{T_1}{T_s} \\ R_i = \dfrac{T_{i+1}}{T_i} \\ K = \dfrac{2T_s}{T_1} = \dfrac{2}{R_s} \end{cases}$$

and noting that

$$R_1 R_2 \ldots R_{n-1} = \prod_{j=1}^{n-1} R_j = \frac{T_n}{T_1}$$

The utilization factor may be written as

$$U = U_s + \sum_{i=1}^{n-1} R_i + \frac{K}{\prod_{j=1}^{n-1} R_j} - n$$

we minimize U over R_i, $i = 1, \ldots, n-1$. Hence,

$$\begin{aligned} \frac{\partial U}{\partial R_i} &= \cancelto{0}{\frac{\partial U_s}{\partial R_i}} - \cancelto{0}{\frac{\partial n}{\partial R_i}} + \cancelto{0}{\frac{\partial \sum_{j\neq i}^{n-1} R_j}{\partial R_i}} + \frac{\partial R_i}{\partial R_i} + \frac{\partial K \left(\prod_{j=1}^{n-1} R_j\right)^{-1}}{\partial R_i} \\ &= 1 + K \left(\prod_{j=i}^{n-1} R_j\right)^{-1} \frac{\partial R_i^{-1}}{\partial R_i} \\ &= 1 - K \left(\prod_{j=i}^{n-1} R_j\right)^{-1} R_i^{-2} \\ &= 1 - \frac{K}{R_i^2 \left(\prod_{j\neq i}^{n-1} R_j\right)} \\ &= 1 - \frac{K}{R_i \left(\prod_{j=1}^{n-1} R_j\right)} \end{aligned}$$

Thus, defining $P = \prod_{j=1}^{n-1} R_j$, U is minimum when:

$$\begin{cases} R_1 P = K \\ R_2 P = K \\ \cdots \\ R_{n-1} P = K \end{cases}$$

that is, when all R_i have the same value:

$$R_1 = R_2 = \cdots = R_{n-1} = K^{1/n}$$

Substituting this value in U we obstain:

$$\begin{aligned}
U_{lub} &= U_s + \sum_{i=1}^{n-1} R_i + \frac{K}{\prod_{j=1}^{n-1} R_j} - n \\
&= U_s + \sum_{i=1}^{n-1} \left(K^{1/n}\right) + \frac{K}{\prod_{j=1}^{n-1} \left(K^{1/n}\right)} - n \\
&= U_s + (n-1)K^{1/n} + \frac{K}{K^{n-1/n}} - n \\
&= U_s + (n-1)K^{1/n} + K\,K^{-n-1/n} - n \\
&= U_s + (n-1)K^{1/n} + K1-n-1/n - n \\
&= U_s + (n-1)K^{1/n} + K^{n-n+1/n} - n \\
&= U_s + (n-1)K^{1/n} + K^{1/n} - n \\
&= U_s + nK^{1/n} - K^{1/n} + K^{1/n} - n \\
&= U_s + n(K^{1/n} - 1)
\end{aligned}$$

Now, noting that

$$U_s = \frac{C_s}{T_s} = \frac{T_1 - T_s}{T_s} = R_s - 1$$

we have

$$R_s = U_s + 1$$

Thus, K can be rewritten as

$$K = \frac{2}{R_s} = \frac{2}{U_s + 1}$$

and finally

$$U_{lub} = U_s + n\left[\left(\frac{2}{U_s+1}\right)^{1/n} - 1\right]$$

Thus, given a set of n periodic tasks and a polling server withg utilization factors U_p and U_s, respectively, the schedulability of the periodic task set is guaranteed under RM if

$$U_p + U_s \leq U_s + n\left(K^{1/n} - 1\right)$$

that is, if

$$U_p \leq n\left[\left(\frac{2}{U_s+1}\right)^{1/n} - 1\right]$$

Appendix C

U_{lub} for RM + Deferrable Server

To simplify the computation of the bound for n periodic tasks, we first determine the worst-case relations among the tasks, and then we derive the lower bound against the worst-case model.

Consider a set n periodic tasks $(\tau_1, \ldots, \tau_n)$, ordered by increasing periods, and a Deferrable Serve with a higher priority. The worst-case condition for the priodic tasks, is such that $T_1 < T_n < 2T_1$. In the presence of a DS, however, the derivation of the worst-case is more complex and requires the analysis of three different cases. For the sake of clarity, here we analyze one case only, the most general, in which DS may execute three times within the period of the highest-priority periodic task. This happens when DS defers its service at the end of its period and lso executes at the beginning of the next period. In this situation, the full processor utilization is achieved by the following tasks' parameters:

$$\begin{cases} C_s = T_1 - (T_s + C_s) = \dfrac{T_1 - T_s}{2} \\ C_1 = T_2 - T_1 \\ C_2 = T_3 - T_2 \\ \cdots \\ C_{n-1} = T_n - T_{n-1} \\ C_n = T_s - C_s - \displaystyle\sum_{i=1}^{n-1} C_i = \dfrac{3T_s + T_1 - 2T_n}{2} \end{cases}$$

Hence, the resulting utilization is:

$$\begin{aligned} U &= \frac{C_s}{T_s} + \frac{C_1}{T_1} + \cdots + \frac{C_n}{T_n} \\ &= U_s + \frac{T_2 - T_1}{T_1} + \cdots + \frac{T_n - T_{n-1}}{T_{n-1}} + \frac{3T_s + T_1 - 2T_n}{2T_n} \\ &= U_s + \frac{T_2}{T_1} + \cdots + \frac{T_n}{T_{n-1}} + \left(\frac{3T_s}{2T_1} + \frac{1}{2}\right)\frac{T_1}{T_n} - n \end{aligned}$$

defining:

$$\begin{cases} R_s = \dfrac{T_1}{T_s} \\ R_i = \dfrac{T_{i+1}}{T_i} \\ K = \dfrac{1}{2}\left(3\dfrac{T_s}{T_1} + 1\right) \end{cases}$$

and noting that

$$R_1 R_2 \ldots R_{n-1} = \prod_{j=1}^{n-1} R_j = \frac{T_n}{T_1}$$

The utilization factor may be written as

$$U = U_s + \sum_{i=1}^{n-1} R_i + \frac{K}{\prod_{j=1}^{n-1} R_j} - n$$

we minimize U over R_i, $i = 1, \ldots, n-1$. Hence,

$$\begin{aligned}
\frac{\partial U}{\partial R_i} &= \cancelto{0}{\frac{\partial U_s}{\partial R_i}} - \cancelto{0}{\frac{\partial n}{\partial R_i}} + \cancelto{0}{\frac{\partial \sum_{j\neq i}^{n-1} R_j}{\partial R_i}} + \frac{\partial R_i}{\partial R_i} + \frac{\partial K \left(\prod_{j=1}^{n-1} R_j\right)^{-1}}{\partial R_i} \\
&= 1 + K \left(\prod_{j=i}^{n-1} R_j\right)^{-1} \frac{\partial R_i^{-1}}{\partial R_i} \\
&= 1 - K \left(\prod_{j=i}^{n-1} R_j\right)^{-1} R_i^{-2} \\
&= 1 - \frac{K}{R_i^2 \left(\prod_{j\neq i}^{n-1} R_j\right)} \\
&= 1 - \frac{K}{R_i \left(\prod_{j=1}^{n-1} R_j\right)}
\end{aligned}$$

Thus, defining $P = \prod_{j=1}^{n-1} R_j$, U is minimum when:

$$\begin{cases} R_1 P = K \\ R_2 P = K \\ \cdots \\ R_{n-1} P = K \end{cases}$$

that is, when all R_i have the same value:

$$R_1 = R_2 = \cdots = R_{n-1} = K^{1/n}$$

Substituting this value in U we obstain:

$$\begin{aligned}
U_{lub} &= U_s + \sum_{i=1}^{n-1} R_i + \frac{K}{\prod_{j=1}^{n-1} R_j} - n \\
&= U_s + \sum_{i=1}^{n-1} \left(K^{1/n}\right) + \frac{K}{\prod_{j=1}^{n-1} \left(K^{1/n}\right)} - n \\
&= U_s + (n-1)K^{1/n} + \frac{K}{K^{n-1/n}} - n \\
&= U_s + (n-1)K^{1/n} + K\,K^{-n-1/n} - n \\
&= U_s + (n-1)K^{1/n} + K1-n-1/n - n \\
&= U_s + (n-1)K^{1/n} + K^{n-n+1/n} - n \\
&= U_s + (n-1)K^{1/n} + K^{1/n} - n \\
&= U_s + nK^{1/n} - K^{1/n} + K^{1/n} - n \\
&= U_s + n(K^{1/n} - 1)
\end{aligned}$$

Now, noting that

$$U_s = \frac{C_s}{T_s} = \frac{T_1 - T_s}{2T_s} = \frac{R_s - 1}{2}$$

we have

$$R_s = 2U_s + 1$$

Thus, K can be rewritten as

$$K = \left(\frac{3}{2R_s} + \frac{1}{2} \right) = \frac{U_s + 2}{2U_s + 1}$$

and finally

$$U_{lub} = U_s + n \left[\left(\frac{U_s + 2}{2U_s + 1} \right)^{1/n} - 1 \right]$$

Thus, given a set of n periodic tasks and a polling server withg utilization factors U_p and U_s, respectively, the schedulability of the periodic task set is guaranteed under RM if

$$U_p + U_s \leq U_s + n \left(K^{1/n} - 1 \right)$$

that is, if

$$U_p \leq n \left[\left(\frac{U_s + 2}{2U_s + 1} \right)^{1/n} - 1 \right]$$

Appendix D

POSIX

D.1 Implementing Periodic Tasks

The pseudocode of a Periodic Task is:

```
1   void *PeriodicTask(void *arg)
2   {
3       <initialization>;
4       <start periodic timer, period = T>;
5       while (condition)
6       {
7           <job body>;
8           <wait next activation>;
9       }
10  }
```

The job body is outside of out control as well as the initialization since it is application dependent. Hence we will analyze the various way we can implement the `<start periodic timer>` and `<wait next activation>` section/function using timers and clocks.

D.1.1 Using UNIX clock

The idea is to use unix clocks to implement the `wait_next_activation` function using `usleep` (relative sleep).

In this naive implementation:

- The program reads the current time
- The program computes the relative sleep

$$\delta = \text{next activation time} - \text{current time}$$

- The program calls `usleep` for a value of δ

```
1   void wait_next_activation(void)
2   {
3       gettimeofday(&tv, NULL);
4       d = nt - (tv.tv_sec * 1000000 + tv.tv_usec);
5       nt += period;
6       usleep(d);
7   }
```

The problem with this implementation is that preemption can happen in `wait_next_activation` between `gettimeofday` and `usleep` resulting in an incorrect sleeping time of the task.

The solution is to call a function that implements a periodic behaviour.

D.1.2 Using UNIX itimer

Unix systems provide a system call for setting up a periodic timer.

```
#include <sys/time.h>
int setitimer (int which, const struct itimerval *value, struct itimerval *
ovalue);
```

The first parameter of the system call is the type of interval timer to use. Three values are admissible:

- `ITIMER_REAL`: timer fires after a specified real time. `SIGALRM` is sent to the process.
- `ITEMER_VIRTUAL`: timer fires after the process consumes a specified amount of time (process time). `SIGPROF` is sent to the process.
- `ITEMER_PROF`: process time + system calls (both user and system time = profiling)

Hence `setitimer()` can be used to implement `<start periodic timer>`.

```
#include <sys/time.h>  // setitimer()
#include <signal.h>    // signal()
#include <unistd.h>    // pause()

#define wait_next_activation pause // pause till a signal is fired

static void sighand(int s) {} // empty signal handler (need for signal())

int start_periodic_timer(uint64_t offs, int period)
{
    struct itimerval t;
    // offset
    t.it_value.tv_sec     = offs / 1000000;
    t.it_value.tv_usec    = offs % 1000000;
    // period
    t.it_interval.tv_sec  = period / 1000000;
    t.it_interval.tv_usec = period % 1000000;
    // register signal and specify signal handler
    signal(SIGALRM, sighand);

    return setitimer(ITIMER_REAL, &t, NULL);

}
```

The problem with this implementation is that the `SIGALRM` generated is handled by an empty handler. An idea is to implement a better `wait_next_activation`: instead of pause, the function can wait a `SIGALRM`.

```
#include <sys/time.h>  // setitimer()
#include <signal.h>    // signal()

static sigset_t sigset;

static void wait_next_activation(void)
{
    int dummy;
    sigwait(&sigset, &dummy); // wait for any signal in the set to be pending
}

int start_periodic_timer(uint64_t offs, int period)
{
    struct itimerval t;
    // offset
    t.it_value.tv_sec     = offs / 1000000;
    t.it_value.tv_usec    = offs % 1000000;
    // period
    t.it_interval.tv_sec  = period / 1000000;
    t.it_interval.tv_usec = period % 1000000;
    // define set of signal that should be processed
    sigemptyset(&sigset);                  // exclude all the defined signals
    sigaddset(&sigset, SIGALRM);           // add SIGALRM to the signal set
    sigprocmask(SIG_BLOCK, &sigset, NULL); // block signals specified in sigset
(required by sigwait)

    return setitimer(ITIMER_REAL, &t, NULL);
}
```

The limitation of using UNIX timers is that only one real-time timer is available per process.

D.1.3 POSIX timers

POSIX offer multiple types of clocks (e.g. `CLOCK_REALTIME` and `CLOCK_MONOTONIC`). In addition, it is possible to set up multiple timers per process (each process can dynamically allocate and start timers). A timer firing generates an asynchronous event which si configurable by the program.

```
#include <time.h> // timer_create()
int timer_create(clockid_t clockid, struct sigevent *sevp, timer_t* timerid);
```

- `clockid` specifies the clock that the new timer uses to measure time.
- `sevp` points to a `sigevent` structure that specifies how the caller should be notified when the timer expires
- `timerid` id of the new timer returned by the system call.

```
#include <time.h> // timer_settime()
int timer_settime(timer_t timerid, int flags, const struct itimerspec*v, struct
itimerspec *ov);
```

- `timerid` timer id
- `flags`: (use 0 = `TIMER_ABSTIME`)
- `v`: interval timer specifications (timer offset `it_value`, timer period `it_interval`)
- `ov`: just set this to NULL... you wont need it trust me

Since we are using POSIX we need to link agains `librt` hence when compiling link the executable using the compiler flag `-lrt`.

With this new timer the implementation of `start_periodic_timer` becomes

```
#include <time.h>      // timer_create(), timer_settime
#include <signal.h>    // signal()
#include <string.h>    // memset()

static sigset_t sigset;

static void wait_next_activation(void)
{
    int dummy;
    sigwait(&sigset, &dummy); // wait for any signal in the set to be pending
}

int start_periodic_timer(uint64_t offs, int period)
{
    struct itimerspec t;
    struct sigevent sigev;
    timer_t timer;
    int res;
    const int signal = SIGALRM;
    // offset
    t.it_value.tv_sec     = offs / 1000000;
    t.it_value.tv_usec    = offs % 1000000;
    // period
    t.it_interval.tv_sec  = period / 1000000;
    t.it_interval.tv_usec = period % 1000000;
    // define set of signal that should be processed
    sigemptyset(&sigset);                   // exclude all the defined signals
    sigaddset(&sigset, SIGALRM);            // add SIGALRM to the signal set
    sigprocmask(SIG_BLOCK, &sigset, NULL); // block signals specified in sigset
(required by sigwait)
    // specify the sigevent
    memset(&sigev, 0, sizeof(struct sigevent)); // empty sigevent structure
    sigev.sigev_notify = SIGEV_SIGNAL;
    sigev.sigev_signo = signal;
    // create timer
    res = timer_create(CLOCK_MONOTONIC, &sigev, &timer);
    // inizialize periodic timer with the specs in t
    return timer_settime(timer, TIMER_ABSTIME, &t, NULL);
}
```

Still a process can still be preempted and the relative sleeping problem is still present in some form. The solution is to use Absolute Time from POSIX timers and clocks

D.1.4 Using POSIX clock and timers with Absolute time

Instead of reading the current time and computing δ based on it `wait_next_activation()` can directly wait for the absolute arrival time of the next job using a `clock_nanosleep()` call.

```
#include <time.h>
int clock_nanosleep(clockid_t clock_id, int flags, const struct timespec *
request, struct timespec *remain);
```

- `clock_id`: specifies the clock agains which the sleep interval is to be measured (`CLOCK_REALTIME`, `CLOCK_MONOTONIC`, `CLOCK_PROCESS_CPUTIME_ID`).
- `flags`: if 0 then specified request is interval, if `TIMER_ABSTIME` then specified request is an absolute time.
- `request` structure specifying time to wait
- `remain` (if not NULL) returns the remaining unslept time

```
#include <time.h>
clock_gettime(clockid_t clk_id, struct timespec* res);
```

- `clock_id`: specifies the clock agains which the sleep interval is to be measured (`CLOCK_REALTIME`, `CLOCK_MONOTONIC`, `CLOCK_PROCESS_CPUTIME_ID`).
- `res` result of the clock reading

```
#include <time.h>      // clock_gettime(), clock_nanosleep()

static struct timespec r;
static int period;

static inline void timespec_add_us(struct timespec *t, uint64_t d){ ... } // add
 d to timespec t by converting to us to ns
static void wait_next_activation(void)
{
    clock_nanosleep(CLOCK_REALTIME, TIMER_ABSTIME, &r, NULL);
    timespec_add_us(&r, period)
}

int start_periodic_timer(uint64_t offs, int period)
{
   clock_gettime(CLOCK_REALTIME, &r);
   timespec_add_us(&r, offs);
   period = t;
   return 0;
}
```

Notice that r and period are global variable: this is a bad idea. Hence we wrap these two variables in a struct:

```
struct periodic_task
{
    struct timespec r;
    int period;
}
```

Whenever a new periodic task is needed: inside `start_periodic_timer` periodic task is initialized and its values are passed to the aforementioned function calls.

Summary

- Unix allows 1 process to have N timers and 1 thread can sleep until a future time point (only relative sleep)
- POSIX allows 1 timer for 1 thread (RT task) and avoids the relative-sleep problem

- POSIX guarantees `CLOCK_REALTIME` exists, but it is not good for real time applications: may jump forward/backward and is prone to overflow: it cannot be restarted hence it will overflow on 32 bit machine in year 2038
- For these reasons, use `CLOCK_MONOTONIC` whenever possible since it is not user settable, it is restarable when the system is restarted (use it in conjunction to `clock_gettime()`)

D.2 Real-Time scheduling

D.2.1 Better Statistics

First and foremost we need to improve the statistics for each task: since now we have considered the average time over 100 jobs as a statistics. However, this measurement may still yield a misleading result.

Example 25: $T = D = 2$

A task τ generates 101 jobs ranging from J_1 to J_{101}.
All but J_2 start at activation time and all but J_1 need 1 time units to complete. J_1 however finishes 3 time units after activation (deadline miss).
The average time printed by J_{101} is 2: 1 job every 2 time units. Through the fact that J_1 missed its deadline caused every other job to miss its deadline.

In other terms we need to introduce additional statistical quantities to verify which job if any miss their deadlines: namely jitter, best and worst case response times and consecutive deadline misses.

Given a job $J_{i,k}$ of a task τ_i we need to save the following quantities:

- k: job number
- $s_{i,k}$: start time of the job
- $f_{i,k}$: finishing time of the job
- t_0: expected time of earliest task activation (initial offset of first job activation)
- For every task τ_i with offset Φ_i, period T_i and relative deadline D_i, compute the absolute activation time and absolute deadline:

$$r_{i,k} = t_0 + \Phi_i + (k-1)T_i \qquad d_{i,k} = r_{i,k} + D_i$$

  Given these quantities we can define:

  – A job $J_{i,k}$ does not miss its deadline if

$$r_{i,k} \leq s_{i,k} \leq f_{i,k} \leq d_{i,k}$$

  – Best and worst case response time

$$\min_k (f_{i,k} - r_{i,k}) \qquad \max_k (f_{i,k} - r_{i,k})$$

  – Relative start time jitter

$$\max_k |(s_{i,k} - r_{i,k}) - (s_{i,k-1} - r_{i,k-1})|$$

  – Absolute start time jitter

$$\max_k |s_{i,k} - r_{i,k}| - \min_k (s_{i,k} - r_{i,k})$$

  – Relative finish time jitter

$$\max_k |(f_{i,k} - r_{i,k}) - (f_{i,k-1} - r_{i,k-1})|$$

  – Absolute finish time jitter

$$\max_k |f_{i,k} - r_{i,k}| - \min_k (f_{i,k} - r_{i,k})$$

These quantities can be measured using the `clock_gettime()` in conjunction with `CLOCK_MONOTONIC`, whereas t_0 can be computed using `timer_settime()` with flag `TIMER_ABSTIME`.
In addition to this the response time $R_{i,k}$ is equal to the execution time $c_{i,k}$ if and only if $r_{i,k}$ up to $f_{i,k}$ has no preemption/blocking, otherwise it holds that $R_{i,k} > c_{i,k}$. To measure the execution time use `clock_gettime()` with `CLOCK_THREAD_CPUTIME_ID`. In addition, since this measurement is affected by energy-saving mode we need to change the cpu scaling for each core in the machine:

```
cd /sys/devices/system/cpu/cpufreq
cat policy?/scaling_{min,cur,max}_freq   # cat scaling_{min,cur,max}_freq content contained in each folder policy$character$
for x in policy?; do
    sudo tee $x/scaling_min_freq < $x/scaling_max_freq; % for each core policy replace scaling_min_freq with scaling_max_freq
done
cat policy?/scaling_{min,cur,max}_freq   # cat scaling_{min,cur,max}_freq content contained in each folder policy$character$
```

D.2.2 Real-Time Scheduling

POSIX provides support for real-time scheduling: but NO multiprocessor parititoning and migration. To provide this support for real time scheduling POSIX provides priority scheduling support:

- Multiple priority levels
- A task queue per priority level
- Run first task fo highest-priority in non-empty queue

POSIX also provides multiple scheduling policies: a scheduling policy describes how tasks are moved between the prioririty queues. Please notice that these scheduling policies consider Fixed priority: task is always in the same priority queue.
Not all RT theory for uniprocessor (UP) is for MP. So we need to partition MP to have 1 processor as 1 UP. In order to do so we add the following api calls:

```
#define _GNU_SOURCE // CPU_ZERO(), CPU_SET(), sched_setaffinity()
#include <sched.h>  // CPU_ZERO(), CPU_SET(), sched_setaffinity()

void main(int argc, char**argv)
{
    cpu_set_t cpumask;
    CPU_ZERO(&cpumask);         // clears cpumask, so that it contains no CPUs
    CPU_SET(0, &cpumask);       // Add CPU 0 to cpumask
    // sets the CPU affinity mask of the thread whose ID is pid (=0),
    // if pid is zero, then the calling thread is used.
    // it also has to provide the size of the cpuset pointed to by cpumask
    if (sched_setaffinity(0, sizeof(cpumask), &cpumask))
    {
        perror("sched_setaffinity failed");
        return -1;
    }
}
```

POSIX specifically requires four scheduling policies:

- `SCHED_FIFO`: tasks with same priority level (in the same queue) are executed in a first in first out manner (this might lead to starvation of other tasks with same priority). Only higher priotiy tasks can preempt it
- `SCHED_RR`: tasks with same priority level (in the same queue) are executed in a round robin fashion: a task executes for a **time/scheduling quantum** (fixed interval of time), than it is suspended and added to the back of the queue (tasks with the same priority are served fairly)
- `SCHED_SPORADIC`: is a sporadic server that decreases the response times of aperiodic RT tasks
- `SCHED_OTHER`: is the traditional Unix scheduler. It has dynamic priorities, and it is scheduled in background with respect to fixed priorities.

time/scheduling quantum

RR and FIFO priority values are comparable:

- `sched_get_priority_min(int $policy$)`: lowest priority value

- `sched_get_priority_max(int $policy$)`: highest priority value

Hence the schduling policy can be set using the following API calls:

```
#include <sched.h>

int sched_get_priority_min(int policy); // minimum priority value that can be used by the policy
int sched_get_priority_max(int policy); // maximum priority value that can be used by the policy

struct sched_param
{
    ...
    int sched_priority;
    ...
}
/*
    @pid:    id of the thread (if 0 calling thread)
    @policy: scheduling policy
    @param:  value between min and max priority value (for SCHED_FIFO and SCHED_RR)
*/
int sched_setscheduler(pid_t pid, int policy, const struct sched_param *param);

/*
    @pid:    id of the thread (if 0 calling thread)
    @param:  value between min and max priority value
*/
int sched_setparam(pid_t pid, const struct sched_param *param);
```

If there is no swap to disk, no unlucky kernel latency, and MP has no unlycky bus contention, the scheduled RT tasks will have no deadline misses.
In general, "regular" (`SCHED_OTHER`) tasks are scheduled in background relative to real-time tasks. A real-time task can preempt/starve other tasks.
Running applications with real-time priroities requires root privileges, hence after compiling perform this bash commands:

```
sudo chown root executable # change file owner to root
sudo chmod u+s executable  # allow root to execute the executable (not the user)
```

D.2.3 Memory Swapping

The virtual memory mechanism can swap part of the process address space to disk: Memory swapping can increase execution times, resulting in temporal unpredictability, hence it is not good for real-time application.

A non-real time task can force a memory swap. To overcome this issue RT task can lock part of its address space in memory:

- Locked memory cannot be swapped out of the physical memory disk
- This can result in a RAM exhaustion

Memory locking can be performed only by applications having root privileges.

```
#include <sys/mman.h>

// lock pages starting at @addr with size len (in bytes)
int mlock(const void *addr, size_t len);
// unlock pages starting at @addr with size len (in bytes)
int munlock(const void *addr, size_t len);
/*
    lock entire address space in memory
    MCL_CURRENT = current address space
    MCL_FUTURE  = all memory allocated in the future
    MCL_CURRENT | MCL_FUTURE = both
*/
int mlockall(int flags);
```

D.3 Concurrency

A process implements the notion of **protection** in the sense that each process has its own address space and other private resources. A process can write/read in its address space but is not allowed to touch other processes' resources: two processes can share some resources for communication, but this has to be excplicitly allowed by them! protection

Processes usually communicate through message passing (e.g. pipes, sockets, signals,...).

A process is more than a set of private resources, it is an active entity. The concept of process involves two aspects:

- **Protection** or resource ownership Protection
- **Execution** Execution
  A process contains at least a schedulable entity, which can access the process's resources. As such it contains all the information for the scheduling (e.g. scheduling parameters). Furthermore this schedulable entity is also characterized by (at least) a CPU state and a stack.

Each single-threaded process has only one thread which involves:

- One address space per process
- One stack per process
- One Process Control Block (PCB) per process
- Other private resources
- One single executino flow per process

This differs from a multi-threaded process in the sense that a process can have multiple threads in it:

- One address space
- One PCB
- Multiple execution flows in a single process
- Multiple stacks (one per thread)
- One Thread Control Block (TCB) per thread

D.3.1 Small Summary about Processes

The Process Memory Layout involves three major components in its private address space: User memory, Stack and Heap. In particular, the user memory is divided into three segments:

1. Initialized data segment
2. BSS (Block Started by Symbol): uninitialized global and static variables
3. Text segment (containing the program code)

The heap and the stack are straightforward to understand. However, the heap is usable through `malloc` and can grow using `brk()` and `sbrk()`

Each process is identified by a process ID (PID) which is unique in the system. Each time a process is created a PID is assigned to it. The value of the PID of a process can be retrieved using `getpid(void)`, whereas the PID of a process's creator (i.e. parent) can be retrieved using `getppid(void)`

D.3.1.1 Fork

A new process can be created using the api call `fork()`. The new process (called child process) contains a copy of the parent's address space. The call has one entry point and two exit points:

- In the child, 0 is returned
- In the parent, the PID of the child is returned

In case of error fork returns a negative value is returned.

Generally speaking `fork()` is used in the following way:

```
#include <sys/types.h> // pid_t
#include <unistd.h>    // fork

pid_t child_pid = fork();
if(child_pid < 0) { return -1 }
if(child_pid == 0)
    // child body
else
    // parent body
```

Alternatively:

```
#include <sys/types.h> // pid_t
#include <unistd.h>    // fork
#include <stdlib.h>    // exit

pid_t child_pid = fork();
if(child_pid < 0) { return -1 }
if(child_pid == 0)
{
    // child body
    exit(0);
}
// parent body
```

Since the child address space is a copy of the parent's: the child's text segment is the same as the parent's, consequently both the parent body and the child body must be in the same executable.

D.3.1.2 Exec

To avoid this one can use the `exec()` call or any of its variants.
Exec is a family of functions to replace the process address space (text, data and heap) (e.g. `execl()`,`execlp()`,`execle()`,`execv()`,`execvp()`). The basic idea of exec is that it loads a new programme and jumps to it, but does not create a new process.

An example usage of exec is as follows:

```
#include <sys/types.h> // pid_t
#include <unistd.h>    // fork
#include <stdio.h>     // perror

pid_t child_pid = fork();
if(child_pid < 0) { return -1 }
if(child_pid == 0)
{
    char *args[3] = {"arg1", "arg2", "arg3"};
    execve("child_body", args, NULL);
    perror("Execve"); // check if any error has occured
    return -1;
}
...
```

Some non-posix compliant systems make no distinction between program and process and only provide a fork + exec combo. POSIX also provides a `system()` function which does fork + exec (+ wait)

A process terminates when:

1. It invokes either the library call `exit()` or the system call `_exit()`
2. It returns from its main function
3. It is killed by some external event (e.g. a signal)

When it terminates explicity, a process can return a result to the parent.
Every process can register a hook ot be called on regular process termination

```
#include <stdlib.h>    // atexit

int atexit(void (*function) (void));
```

However, please note that handlers are not called if exiting with `_exit()`.

D.3.1.3 Wait

First form of synchronization between processes is waiting for a children to return to the parent.
A parent can wait for its child's termination using the api calls `wait()`, `waitpid()`, `wait4()`.
Formally:

```
#include <sys/types.h>     // pid_t
#include <sys/wait.h>      // wait()

pid_t wait(int *status);
```

If the calling process has no children, `wait()` returns a value < 0
If at least there is one terminated child, `wait()` returns the child's exit value, and child's private resources are freed
If there are no terminated children, `wait()` blocks the calling process.

The other variants of the wait family of function allow to select the child to wait for.

Please note that after a process terminates, its private resources are not freed until its parent performs a `wait()`. Children processes that have terminated are said to be in a **zombie** state. `zombie`
Hence a good parent has to wait for its children to terminate. When the parent of a process dies the process is reparented to `init` (a system process whose PID is 1). As a consequence, when a process dies all its zombies are eliminated.

A process can be notified about the termination of a child process through an asynchronous event (e.g. signal: `SIGCHLD`)

D.3.2 synchronization through Signals

Concurrent processes interact in different ways: competition or cooperation.
Cooperation can be implemented through **signals**: Sometimes, a process has to wait until cooperating processes have completed some operation. In other terms, a process τ_i waits for an asynchronous event generated by: another process τ_j or the system. `signals`

Definition 39: Signal

Asynchronous event directed to process τ. Process τ can:

- Wait for a signal
- Perform some other work in the meanwhile, and the signal will interrupt it.

Signals are the software equivalent of interrupts. A process receiving a signal can:

- Ignore it
- Interrupt its execution and jump to a **signal handler** `signal handler`
- Abort

A signal that has not caused one of the previous actions is a **pending signal**. `pending signal`

Signals are managed using a **signal table**. Such structure is a per process private resource that specifies how the process handles each signal. At process creation the signal table is initialized using default values. `signal table`

The table entries can be modified using: `signal()` or `sigaction()`.

```
#include <signal.h>
/*
    @signum signal (e.g. SIGALRM)
    @handler either SIG_IGN, SIG_DFL or the address of a programmer-defined
function
*/

sighandler_t signal(int signum, sighandler_t handler);
/*
    @signum signal (e.g. SIGALRM) excluding SIGKILL and SIGSTOP
    @act    the new action for signal signum
    @oldact old action for signal signum (if not null)
*/
int sigaction(int signum, const struct sigaction *act, struct sigaction *oldact)
;
```

The prototype of a signal handler is very simple: it must return void and accept an integer value as input `void sighand(int n)`. In order to set up a signal handler, we need to provide some of the parameters in a structure `sigaction` provided by the api.

```
struct sigaction {
    void     (*sa_handler)(int);
    void     (*sa_sigaction)(int, siginfo_t *, void *);
    sigset_t   sa_mask;
    int        sa_flags;
    void     (*sa_restorer)(void);
};
```

where:

- `sa_handler` is the signal handler (function), or `SIG_DFL` (default action) or `SIG_IGN` (ignore the signal)
- `sa_mask` is a mask of signals to disable when the handler runs. This set can be modified using `sigemptyset()`, `sigfillset()`, `sigaddset()`, and `sigdelset()`
- Running handler cannot be interrupted by the handled signal again unless `SA_NODEFER` is set in `sa_flags`

A process can send a signal to other processes by using

```
    int kill(pid_t pid, int sig)
```

Note that it must have the proper permissions: user root can send signals to every process, other users can send signals only to their own processes. `kill` is not used only to kill a process: `kill -SIGALRM` does not kill a process, but `kill -SIGKILL` does.

A process can also send a signal to itself by

```
    int raise(int sig);
```

D.3.2.1 Signal Numbers

Signals are identified by numbers and macros:

- `SIGUSR1` and `SIGUSR1` are user defined signals
- `SIGALRM`, `SIGVTALRM`, and `SIGPROF` are used by process real-time, virtual and profiling timers
- `SIGKILL` is used to kill a program
- `SIGCHLD` is raised every time a child dies, or it is stopped, or it is resumed
- `SIGSTOP`, `SIGTSTP`, `SIGTTIN`, `SIGTTOU` stops a child
- `SIGCONT` resumes a child process

These last three signal numbers are useful for avoiding zombies:

- The `SIGCHLD` handler can perform a `wait()`
- If `SIGCHLD` is ignored, zombies are not created
- Suppress `SIGCHLD` for stopped/resumed children by setting `SA_NOCHLDSTOP` in `sa_flags`

The main problem with signals are as follows:

- Almost all of the signals are reserved for the system, only `SIGUSR` are free for user programs
- Signals can be lost: if a signal arrives more than once while it is blocked, it is not queued (it will fire only one time), this makes signals quite unreliable for RT IPC
- Signals do not transport information, only the signal number is available to the handler

The solution is to use POSIX Real-Time signals

D.3.3 Real-Time signals

Multiple pending RT signals are queued, not lost: the unblocked signal with a lower signo are delivered first. POSIX specifies no further ordering but Linux does.

Signals can transport information in the form of integers or pointers. For this reason the signal handler function prototype has an extended form:

```
void sighand(int signum, siginfo_t *info, void *ignored);
```

As before such handler can be set using `sigaction`, with the only difference that:

- instead of using `sa_sighandler`, you must use `sa_sigaction`
- You must use `SA_SIGINFO` in `sa_flags`.

This allows us to use a $M = SIGRTMAX - SIGRTMIN + 1$ signals for user app. Instead of using `kill()` to send a message, now you should queue all the RT signal using `sigqueue()`

Real-Time signals carry information in `siginfo_t`

```
typedef struct
{
    int si_signo;          // signal number
    int si_code;           // cause of the signal
                           // kill() => SI_USER,
                           // sigqueue() => SI_QUEUE
                           // posix timer => SI_TIMER
    union sigval si_value; // information carried by the signal
} siginfo_t;

union sigval
{
    int sival_int;
    void *sival_ptr;
}
```

The signal is queued as follows:

```
#include <signal.h>

int sigqueue(pid_t p, int n, const union sigval value);
```

`sigqueue()` returns < 0 in case of error, `EAGAIN` is returned if queue is full (no lost signal).

If no error occurs, queue signal n to process p and the information *value* is transmitted with the signal.

The OS can generate RT signals, e.g., when POSIX async I/O is used to impplement future and promise.

For this reason we must specify the RT signal the OS shall generate by using the structure:

```
struct sigevent
{
    int sigev_notify;          // SIGEV_SIGNAL
    int sigev_signo;           // set this to the RT signal
    union sigval sigev_value; // information
    void (*sigev_notify_function)(union sigval);
    void *sigev_notify_attributes;
    pid_t sigev_notify_thread_id;
};
```

POSIX's System Interfaces - Signal Actions says:

- If the process is multi-threaded or
- If the process is single-threaded and a signal handler is executed other than as the result of either:
  - The process calling `abort`, `raise`, `kill`, `pthread_kill` or `sigqueue` to general a signal that is not blocked, or
  - A pending signal being unblocked and being delivered before the call that unblocked it returns.

Then any variable of static storage duration (global variable) used by signal handler should be limited to:

- `errno` (if used, save and restore at start and finish)

- write (never read) `volatile sig_atomic_t` vars

And, any function called by handler should be limited to async-signal-safe function.

D.4 POSIX Thread and their Real-Time Scheduling

The POSIX standard is an IEEE standard specifying the OS interface. As such it is implemented by most Unix systems. The core concept of POSIX is that it defines a C API to handle concurrent activities and it distinguishes between processes and threads.
A thread is a schedulable entity (a flow of execution).
A process has one or more threads+ some private resources (address space, file table, ...). So one thread is a single flow of control withing a process whereas a process has at least one thread (i.e. the main thread).

D.4.1 Threads

All threads in a process share the same address space, file table, program text, ... Each thread has it own context and its own stack. In each process there is a "special" thread (the main thread), which termination causes the termination of the process.

Therefore the question is: should we use processes or threads?
From a general point of view they both are an abstraction for parallelism, but processes are a protection boundary, whereas threads do not. This means that a process sees no other process. However, IPC is usually costly for RT application and unsupportive of RT communication (e.g. Priority inheritance).

On the other hand a thread is not a protection boundary, in the sense that a thread can see other threads and their data (states). As such other threads can screw up another thread. However, threads do not have costrly IPC because communication is done by shared data. In addtion, threads support RT comunication such as Priority Inheritance and Highest Locking Priority protocols.

POSIX defines itw own threading library called `pthread` library which implements all primitive operations on threads (e.g. creation, termination, synchronization, ...). Threading primitives and data structures are contained in their separate headers (e.g. `sched.h`,`pthread.h`,`semaphore.h`). Furthermore, in order to use this headers we must compire the program using the compiler flag `-pthread`.

The code executed by a thread (i.e. the **thread body**) is a C function of the form: thread body

```
1    void *thread_body(void *arg) { ... }
```

When created, a thread starts executing the first instruction of its body. The thread ends when exiting the body (at the end of the C function), but a thread can terminate also in other ways, such as by explicitly calling a termination function, when killed by another thread, ...

A thread is created by invoking `pthread_create()`

```
1    #include <pthread.h>
2    /*
3        @tid the created thread ID
4        @attr specifies some thread's attributes
5        @body is a pointer to the thread body
6        @args is passed to the thread body upon execution
7        @return 0 if no error occurred
8    */
9    int *pthread_create(pthread_t *tid, pthread_attr_t *attr,
10   void*(*body)(void*),void *arg);
```

Each thread is identified by a unique thread ID (TID). The current thread TID can be obtained by invoking

```
1    #include <pthread.h>
2    pthread_t *pthread_self(void);
```

Two TIDs can be compared by invoking:

```
1    #include <pthread.h>
2    int *pthread_equal(pthread_t id1, pthread_t id2);
```

thread attributes specified in `attr` control some characteristics of the created threads such as: stack size (and address), detach state (joinable or detached), some scheduling parameters (priority,...).
Thread attributes must be initialized and destroyed:

```
#include <pthread.h>
int pthread_attr_init(pthread_attr_t *attr);
int pthread_attr_destroy(pthread_attr_t *attr);
```

Thread can terminate by invoking `pthread_exit()`

```
#include <pthread.h>
void pthread_exit(void *retval);
```

A thread can also terminate when its body returns: if the thread is not the main thread `pthread_exit()` is automatically called, if the thread is the main thread `exit()` is called.
Of course if the main thread terminates, the process terminates.

Thread can wait for another thread's termiantion by

```
#include <pthread.h>
int pthread_join(pthread_t id, void **result);
```

Terminated thread's return value is returned in `result` (`*result == PTHREAD_CANCELED` if the thread was killed)

Every thread should be joined:

- The private resources of a terminated thread are not freed until joined (think about memory leak)
- Similar to `wait()` on child processes (think about zombies)

A thread that won't be joined must be detached: when a detached thread terminates its resources are immediately released. There are two ways to detach a thread:

- Detach it at creation by `attr` parameter. Setting `attr` is performed using `pthread_attr_setdetachstate()`
- Detach it after creation by calling `pthread_detach()`

Joining a detached thread results in an error.

A thread terminates by:

- Returning from the body
- Calling `pthread_exit()`
- Killed by other threads using

```
#include <pthread.h>
int pthread_cancel(pthread_t tid);

```

`tid` is that of the thread to be killed. The function returns 0 if and only if no error occurs.

Terminated thread's private resources are released: when the thread is joined (if it is not detached), immediately (if the thread is detached).

Sometimes, killing a thread can leave the system in an inconsistent state (e.g. when settings HW device), so thread cancellation can be deferred until the system state is consistent.

Thread's cancellation si determined by its **cancellability state**, consisting of a cencellability type and enable/disable flag. The cancellability type can be: `cancellability state`

- **Deferred** (default): execute thread cancellation request only at cancellation points, so the thread is canceled only when its state is consisten `Deferred`
- **Asynchronous**: execute thread cancellation request immediately `Asynchronous`

In order to set the cancellability state's type one can use:

```
#include <pthread.h>
/*
    PTHREAD_CANCEL_DEFERRED
    PTHREAD_CANCEL_ASYNCHRONOUS
```

```
    */
    int pthread_setcanceltype(int type, int *oldtype)
```

Whereas the cencallability state's enable/disable is set using:

```
    #include <pthread.h>
    /*
        PTHREAD_CANCEL_ENABLE
        PTHREAD_CANCEL_DISABLE // All cancellation requests, regardless of type, are
     queued until cancellability is enabled again
    */
    int pthread_setcancelstate(int state, int *oldstate)
```

With `PTHREAD_CANCEL_DEFERRED`, cancellation request terminates thread only at cancellation point. So a cancellation point checks if any cancellation request is pending. A cancellation point can be:

- Explicit if thread calls `pthread_testcancel()`
- Implicit if thread calls: any standard I/O function, `pthread_cond_wait()`, ...

When using `PTHREAD_CANCEL_ASYNCHRONOUS` cancellation request ends thread immediately. When using `pthread_cond_wait()`, pending cancellation request ends thread without releasing the waited cond. Hence, in order to leave the system in a consistent way use **cancellation handlers** aka **cleanup handlers**.

cancellation handler

A cleanup handler is a C function with prototype `void handler(void *)`.

cleanup handler

It is called when a thread ends (even if it is killed) and it is the last chance to leave everything in order.

Multiple cleanup handlers can be set: the handlers are stacked and called in a LIFO order. In order to stack handler call:

```
    #include <pthread.h>
    // @arg is passed to handler upon invocation
    void pthread_cleanup_push(void (*handler)(void*), void *arg);
```

To remove an handler call

```
    #include <pthread.h>
    // @execute if not equal to 0, popped handler is executed
    void pthread_cleanup_pop(int execute);
```

`pthread_cleanup_push()` and `pthread_cleanup_pop()` must be paired lexically (e.g. malloc and free)

Thread can set its scheduling policy and parameters by specifying them in `pthread_create()`'s `attr`. This is done using the functions:

```
    #include <pthread.h>
    int pthread_attr_setschedpolicy(pthread_attr_t *attr, int policy);
    int pthread_attr_setschedparam(pthread_attr_t *attr, const struct sched_param *
    param);
```

The only field `sched_priority` matters in `struct sched_param`. In addition the attribute:

```
    #include <pthread.h>
    /*
        PTHREAD_INHERIT_SCHED  inherit creating thread's policy and parameters
        PTHREAD_EXPLICIT_SCHED
    */
    int pthread_attr_setinheritsched(pthread_attr_t *attr, int inheritsched);
```

Default `inheritsched` cannot be assumed (it is implementation-dependent)

D.5 Thread Synchronization

Threads avoid IPC by sharing private resources, trading reliability for speed. So the natural way to synchronize threads is by using the shared resources paradigm.
In particular, there can be two kind of interactions between threads in one process:

- Cooperation: when multiple threads need to synchronize to provide one service, such as mailbox, pipeline, ...
  A complex algorithm can be parallelized by splitting it into a set of parallel activities. This is done to simplify programming and to benefit from multi-/many-core architecture.
  Every parallel activity is executed in a thread that: works on data produced by another thread, produces data for another thread or works on different parts of shared data structure.
  Data/thread not ready for processing/communication can wait by blocking till data/thread's ready, upon which the waiting thread will be unblocked. Cooperation che be achieved using a synchronization mechanism.
- Competition: when different threads use a shared resource that can be used only by one thread at a time.
  A shared resource is usable by 1 thread at a time. So, it must be accessed in **mutual exclusion** (mutex). The code accessing the shared resource is called critical section: only 1 thread shall execute in the critical section. Competition can be achieved using a synchronization mechanism.

`mutual exclusion`

D.5.1 Posix Condition variables vs Mutex

Definition 40: Condition variable

A condition variable is not a variable in the sense that it does not store a value. Multiple threads can block on the variable, waiting for a signal to arrive. Multiple threads can signal the variable:

- If no thread blocks the variable, the signal is lost
- If one or more already block the variable, the signal unblocks one or all the threads blocking it

Definition 41: Mutex

Object with two states (lock and unlock).
Multiple threads can lock a mutex:

- Already unlocked: 1 thread is selected to own the mutex and the rest blocks on the mutex
- Already locked: they block on the mutex

Only mutex's owner can unlock a mutex.

Please notice that a condition variable needs a mutex: if a condition variable c is used without mutex, the signal can be lost incorrectly

- B waits on c but is preempted before completing the wait by queuing itself on c
- A sends signal s to c. s hits empty queue
- A is preempted after completing the sent
- B resumes by queueing itself to c
- B lost s, but B already waits before A sends s

So operations on c can be protected by a mutex m.

When a thread A signals a condition variable c, A shall already own the associated mutex m. The signal then hits the queue of c. If thread B in the queue is chosed for unblocking, some unblocking semantics are possible:

- B is dequeued from c and tries to lock m such that B is guaranteed to own m once A unlocks m
- B is dequeued from c, and ownership of m is transferred from A to B, blocking A on m
- B is dequeued from c and tries to lock m without guarantee of ownership once A unlocks m

A POSIX condition variable si a variable of type `pthread_cond_t`. It must be initialized by

`PTHREAD_COND_INITIALIZER` or

```
#include <pthread.h>
/*
   @attr: NULL for default
   return 0 if and only if init is successful
*/
int pthread_cond_init(pthread_cond_t *cond, const pthread_condattr_t *attr);
```

The condition variable is waited by (analogous to `sigsuspend` in IPC)

```
#include <pthread.h>
/*
   @mutex: mutex already owned by the caller
*/
int pthread_cond_wait(pthread_cond_t *cond, pthread_mutex_t *mutex);
```

Upon return, caller still owns `mutex`

It is signaled by (`mutex` should already be owned)

```
#include <pthread.h>
int pthread_cond_signal(pthread_cond_t *cond); // Unblock 1
int pthread_cond_broadcast(pthread_cond_t *cond);  // Unblock all
```

It is freed by:

```
#include <pthread.h>
int pthread_cond_destroy(pthread_cond_t *cond);
```

If N threads are waiting for a signal, once signaled:

- Scheduling policy determines dequeuing order
- Predictable in a uniprocessor system
- Not automatically so in a multiprocessor system (but can be made predictable with some schemes)
- Dequeue thread(s) synchronously but may run later
- Valid without owning related mutex

POSIX allows `pthread_cond_signal()` to err:

- Sometimes 2 or more threads can be unblocked
- Known as spurious wakeup

D.5.2 Programming practice to avoid deadlock

Settings:

- ≥ 2 shared resources with mutually exclusive access $S_1, \ldots, S_n$
- ≥ 2 threads want to own $S_1, \ldots, S_n$ one after another (e.g. to update them atomically)

Problem: deadlock

- A owns S_1 and is about to lock S_2
- B owns S_2 and is about to lock S_1
- A blocks waiting for S_2 forever
- B blocks waiting for S_1 forever

Solution: Every thread must lock $S_1, \ldots, S_n$ in one particular order

- A owns S_1 and is about to lock S_2
- B is about to lock S_1
- A owns S_2, and B blocks waiting for S_1

The main problem with conditional variable is that `pthread_cond_wait()` is a cancellation point whereas `pthread_mutex_lock` is not. If a thread is cancelled while blockign after invoking `pthread_cond_wait()`, its mutex is owned again before terminating, resulting in the thread dying

while owning the mutex: as a consequence other threads using it will block forever. The solution is to use a cleanup handler to unlock the mutex.

Available RT resource sharing protocols:

- Priority ceiling (`PTHREAD_PRIO_PROTECT`)
- Priority inheritance (`PTHREAD_PRIO_INHERIT`)

Please note that not all implementations support them.

Specify it in `pthread_mutex_init()`'s `attr` by:

```
#include <pthread.h>
int pthread_mutexattr_setprotocol(pthread_mutexattr_t *attr, int protocol);
```

For priority ceiling, specify ceiling priority by:

```
#include <pthread.h>
int pthread_mutexattr_setpriorityceiling(pthread_mutexattr_t *attr, int
prioityceiling);
```