Abstract

Mixed-criticality systems combine real-time components of different levels of criticality, i.e. severity of failure, on the same processor, in order to obtain good resource utilisation. They must guarantee deadlines of highly-critical tasks at the expense of lower-criticality ones in the case of overload. Present operating systems provide inadequate support for this kind of system, which is of growing importance in avionics and other verticals. We present an approach that provides the required asymmetric integrity and its implementation in the high-assurance seL4 microkernel.

1 Introduction

Traditionally, critical real-time systems use dedicated microcontrollers for each function. With increasing functionality and complexity of cyber-physical and other real-time systems, this is creating space, weight and power (SWaP) problems, which force consolidation onto a smaller number of more powerful processors. For example, top-end cars reached 100 processors a few years ago [Hergenhan and Heiser, 2008]; with the robust packaging and wiring required for vehicle electronics, the SWaP problem is obvious, and a driver for the adoption of multitasking OSes [AUT, 2015].

The potential for consolidation is limited unless it is possible to safely co-host functions of different criticality, where criticality is a well-established notion that represents the severity of failure [RTCA]. Certification standards require that safe operation of a particular component must not depend on any less-critical components [ARINC].

Such mixed-criticality systems (MCS) are becoming the norm in avionics, but presently in a very restricted form: the system is orthogonally portioned spatially and temporally, and partitions are scheduled round-robin with fixed time slices [ARINC]. This limits integration and cross-partition communication, and implies long interrupt latencies and poor resource utilisation. The simple partitioning approach will not meet the requirements of future mixed-criticality systems [Barhorst et al., 2009].

Fundamental to good resource utilisation in MCS is the ability to over-commit safely: The system’s core integrity property is that deadlines of the highest criticality tasks must be guaranteed, meaning that there is always time to let such tasks execute their full worst-case execution time (WCET). This may be orders of magnitude larger than the typical execution time, and computation of safe WCET bounds for non-trivial software tends to be highly pessimistic [Wilhelm et al., 2008]. This means that most of the time the highly-critical components leave plenty of slack, which should be available to less critical components, but must be available to the critical component when needed.

Such a system needs support for downgrading timeliness guarantees selectively, least critical ones first. In general, this cannot be achieved by simply giving the most critical tasks the highest priority. Consider the simplified architecture of an autonomous aerial vehicle (AAV) in Figure 1. The most critical component is the low-level flight control, which keeps the vehicle stable and moving towards a waypoint. It executes every 100 ms and normally takes about 10 ms but has a WCET of 70 ms. Next critical are the mission plan, sensor filtering and C&C components, which have execution
rates of between 1 and 10 Hz, normally run for a combined 200 ms every second but have a combined WCET of 500 ms per second. The CAN bus, which connects a video camera and various sensors of secondary importance, can deliver packets every 12.5 µs and does not buffer.

If the critical components are given higher priority than the CAN driver, it will drop many packets even during normal operation, despite the system having sufficient headroom to run everything. The standard realtime (RT) scheduling approach is rate-monotonic priority assignment (RMPA) [Liu and Layland, 1973], which gives highest priority to tasks with the shortest periods. RMPA is easy to analyse and known to be optimal for fixed priorities; it is highly desirable to retain it for MCS.

A further complication is that components of different criticality must be able to communicate, and access shared data [Burns and Baruah, 2013]. For example, the AAV’s mission plan defines the waypoints to be used by the flight control, including some fail-safe return-home path in case the AAV loses ground-station connectivity. It is updated by the ground station via the command and control (C&C) component, and amended by the sensor filtering component for obstacle avoidance; the latter component receives input from various sensors, including camera and other sensor input via the CAN bus. Such communication, including concurrency control between components accessing the same data, must be possible while guaranteeing critical deadlines.

In summary, an OS for mixed-criticality systems must:

- provide high-assurance spatial and temporal isolation, to allow critical components to be assured independently of less critical ones;
- decouple criticality from priority, to ensure critical, low-rate threads meet their deadlines;
- provide mechanisms that allow analysing the timeliness of critical tasks, even if they communicate with less critical ones;
- have well-understood temporal behaviour, especially bounded and known WCET for all operations;
- be highly assured for correct operation.

No such OS exists to date. We present the design and implementation of such an OS, based on the seL4 microkernel for single-core systems. seL4 is an attractive starting point, as it is a high-assurance OS kernel that has been comprehensively verified [Klein et al., 2014], and is the first and still only protected-mode OS in the literature with a complete and sound WCET analysis [Blackham et al., 2011].

We do not claim to have invented new scheduling models or theory. In fact, the system we present in Section 3 is, as scheduling theory goes, known as static mixed criticality [Baruah et al., 2011]. Our claims are about practical systems, specifically:

1. the design of a low-overhead temporal resource management model that is based on a small number of simple, policy-free mechanisms, suitable for a high-assurance implementation, matches the above requirements of MCS but also supports a wide range of other uses (Section 3);
2. its implementation in the seL4 microkernel in a way that retains seL4’s general-purpose nature and verifiability (Section 4);
3. an evaluation that demonstrates that the modifications do not unduly impact seL4’s performance, and support low-overhead implementations of different real-time and best-effort scheduling models (Section 5).

2 Background and Related Work

In the rest of this paper, and this section specifically, we talk about general real-time concepts as well as OS abstraction. Specifically there are two related concepts relating to the execution model. We will use the term task in the sense established in the RT community, namely a set of related jobs which jointly provide some system function, where a job is a unit of work that is scheduled and executed by the system [Liu, 2000]. We use the term (kernel-scheduled) thread to refer to the execution abstraction familiar to the OS community. The term “job”, which we will not use further, corresponds to a unit of work that is conducted by a thread, while “task” maps onto a thread, plus code and data.

In short, we will use “task” when referring to general RT issues, and “thread” when talking about a specific OS concept. In practice, the terms are largely interchangeable.

2.1 Scheduling models vs. mixed criticality

RT scheduling generally assumes periodic tasks, which maps well onto typical control systems, where different activities execute periodically albeit with different periods. Non-periodic (“sporadic”, i.e. interrupt-driven) tasks are incorporated in such a model by requiring a defined minimum arrival time, corresponding to a maximum interrupt rate, which is used as the task’s period for the schedulability analysis. RT tasks have a deadline by which a computation must be finished. The general assumption is that deadlines are implicit, meaning the deadline is the end of the period.

As discussed in the introduction, the ability to overload, while guaranteeing critical deadlines, is core to the notion of MCS. Classical RT scheduling approaches

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1We have not formally re-verified the modified kernel, and only claim that our modifications are moderate in terms of kernel changes and no more difficult to verify than the baseline kernel.
have a notion of (fixed or dynamic) priority as the sole determinant of access to CPU time, with equal-priority tasks (if permitted) being (preemptively or non-preemptively) scheduled FIFO. If the system is overloaded, this means that the lowest-priority deadlines are missed. When using RMPA, this victimises the tasks with the lowest rates. In effect, criticality equals rate in RMPA.

The main alternative to RMPA is earliest deadline first (EDF) scheduling. This is a dynamic priority scheme, which at any time schedules the task with the closest deadline. Unlike fixed-priority schemes, such as RMPA, EDF is optimal on a uniprocessor in that it can schedule any task set, as long as the total utilisation does not exceed 100%. However, the dynamic prioritising implies that under overload, EDF drops deadlines of all tasks [Buttazzo 2005], meaning that there is no concept of task criticality at all.

MCS require control over which deadlines will miss in the case of overload: those of the tasks with low criticality (called LOW tasks from now on), while guaranteeing deadlines of HIGH tasks, irrespective of scheduling priority. This requires a mechanism for limiting CPU time of high-priority tasks.

An established way of providing isolation is through scheduling reservations [Mercer et al. 1993 Oikawa and Rajkumar 1998], where a reservation guarantees a certain share of the CPU to a periodic task. Such schemes are popular in soft RT systems, e.g. multimedia, and some allow slack time to be used by best-effort tasks [Brandt et al. 2003]. Scheduling reservations can be implemented as sporadic servers for RMPA [Sprunt et al. 1989] and with constant bandwidth servers (CBS) [Abeni and Buttazzo 2004] on EDF.

Reservations present a guarantee by the kernel that the reserved bandwidth is available. This means that they do not support over-committing. Also, the kernel must perform a schedulability analysis as admission control whenever a reservation is created. Schedulability tests can be complicated and frequently constitute a trade-off between cost of the test and achievable utilisation.

Recently the concept of a mode switch was introduced to support mixed criticality [Burns and Davis 2014]: when the system is unable to meet its deadline, it enters a high-criticality mode, where the priority of HIGH tasks is boosted above all LOW tasks. To achieve this, HIGH tasks are assigned multiple reservations, one per criticality level. In a two-criticality system, HIGH tasks have a pessimistic WCET and an optimistic worst-observed execution time (WOET). LOW tasks have just one estimate. When the system is in LOW mode, HIGH tasks run according to their WOET. If all reservations in this mode are schedulable, temporal isolation is guaranteed. However, if a HIGH task exceeds its WOET, the system switches to HIGH mode, degrading LOW tasks and assuring asymmetric protection between HIGH and LOW threads without falsely correlating rate and urgency.

### 2.2 Support for sharing and communication

As indicated in the introduction, integrity of critical components must be assured even when tasks communicate and share. In our AAV example of Figure 1, the mission plan component encapsulates waypoints. The HIGH flight-control component must be able to access a consistent view of the flight plan, despite the LOWER C&C and other components performing updates.

Encapsulating the shared data and the code that accesses and modifies it into a single-threaded resource server [Brandenburg 2014] is a simple and effective way to achieve the necessary transaction semantics. Obviously, this server has the criticality level of its most critical client, but must also act on behalf of a LOW client. This creates a temporary criticality inversion where the LOW client blocks the HIGH one. This is an unavoidable consequence of sharing, and the design must ensure that it does not cause the HIGH task to miss deadlines.

![Figure 2: Comparison of real-time locking protocols](image)

**Figure 2:** Comparison of real-time locking protocols based on implementation complexity and priority inversion bound.

There are multiple ways to achieve mutual exclusion in fixed-priority RT systems [Sha et al. 1990], the most common being non-preemptive critical sections (NCP), the priority inheritance protocol (PIP), and the immediate and original priority ceiling protocols (IPCP and OPCP). For RT systems, the most important factor for mutual exclusion is the bound on priority inversion, where a low priority task blocks a high one. For efficient systems, the concern is execution cache performance, for secure systems the concern is the avoidance of channels. Figure 2 shows the four protocols in terms of complexity and priority inversion, none is a silver bullet. NCP is simplest yet has the longest blocking time, IPCP requires the priorities of all lockers to be known a priori. PIP has high implementation complexity and risks deadlock if resource ordering is not used. OPCP is even

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2 Also known as highest lockers protocol and PRIO_PROTECT in POSIX.
more complex, and requires global state to be maintained across all locks in the system, which is not acceptable for seL4 as it introduces covert channels and is incompatible with seL4’s decentralised user-level resource management. We will show in Section 3.2 how IPCP can be easily implemented without the kernel requiring knowledge about critical sections.

As a mechanism for supporting sharing, Fiasco [Steinberg, 2004] introduced the idea of scheduling contexts, separate to execution contexts (threads). Scheduling contexts encapsulate priority, scheduling parameters and accounting detail, and pass between threads over IPC. Steinberg et al. [2010] extended this with bandwidth inheritance [Lamastra et al., 2001; Lipari et al., 2004; Paggioli et al., 2010] over IPC. This is equivalent to PIP combined with reservations. When an IPC from client B arrives at a server S, who is serving a client A with an expired budget, B budget is used to complete A’s request such that B does not have to wait for A’s reservation to be replenished. This kernel-implemented policy, also referred to as helping, prevents the server from choosing alternatives which might be more appropriate in a particular situation, such as aborting A’s request.

The Fiasco design of scheduling contexts [Lackorzynski et al., 2012] is tied to the traditional L4 model of sending IPC messages directly to threads, a model which has been abandoned in modern L4 kernels (including Fiasco and seL4) as it introduces covert channels [Shapiro, 2003]. It is not supported on the later, capability-based Fiasco.OC kernel.

Composite [Parmer and West, 2008] completely frees the kernel from any scheduling policy by providing mechanisms for hierarchical user-level scheduling. It reduces overhead-related capacity loss by configuration buffers shared between user-level and the kernel. Some capacity loss remains as timer interrupts must be delivered down the scheduling hierarchy. This approach does not suit seL4, as the required reasoning about concurrent access (by kernel and user-level) to those buffers would drastically increase verification overhead [Klein et al., 2014].

Unlike all L4 microkernels, Composite implements a migrating thread model [Ford and Leprau, 1999]. This implies that access to shared resources does not block, thus avoiding priority inversion, although at the cost of requiring all server code to be re-entrant, which is fairly heavy-handed policy for a microkernel. Also, it only shifts the problem, as mutual exclusion is still needed, including a way of limiting priority inversions. Given the challenges of getting concurrent code right, it should be minimised in high-assurance systems.

Linux introduced an implementation of the POSIX SCHED_DEADLINE in 3.14, which implements EDF with CBS for temporal isolation. However RT tasks in Linux are higher priority than all other tasks in the system, and cannot be over-committed (although cgroups allow limiting the RT class to a certain share of the CPU). Quest-V [Li et al., 2014] and PikeOS [Kaiser and Wagner, 2007] are both separation kernels for multicore systems that dedicate cores to different criticalities. AUTOBEST [Zuepke et al., 2015] is another separation kernel where the authors demonstrate implementations of AUTOSAR and ARINC653 in separate partitions.

### 2.3 seL4

seL4 is a high-performance OS microkernel with an unprecedented degree of assurance: it features formal proofs of implementation correctness down to the binary, proofs of spatial isolation properties (enforcement of confidentiality, availability and integrity) and a complete and sound analysis of worst-case execution times on ARMv6 processors [Klein et al., 2014]. This assurance makes seL4 an appealing candidate OS for critical systems.

seL4 is designed to be a general-purpose platform, supporting a wide range of use cases. This is a reason why it has a strong emphasis on performance, as many of the envisioned deployment scenarios are performance-sensitive (e.g. mobile devices). Formal verification is a strong motivator for generality: the cost of assurance is best amortised if all use cases are supported by the same, unmodified kernel [Heiser and Elphinstone, 2016]. As the maintainers commit to re-verify any changes to the mainline kernel, they are only interested in changes that make the kernel more general, not more specialised.

#### 2.3.1 seL4 overview

In line with the microkernel minimality principle [Liedtke, 1995], seL4 only provides a small number of policy-free mechanisms. Specifically it provides for threads, represented as thread control blocks (TCBs), address spaces, which are thin wrappers around hardware page tables, and frame objects, which represent physical memory that can be used to populate address spaces by mapping. It further provides port-like endpoint objects for synchronous (rendezvous-style) communication and notification objects, which are essentially arrays of binary semaphores.

Like other security-oriented systems, seL4 uses capabilities [Dennis and Van Horn, 1966] for controlling access to all spatial resources and providing complete mediation similar to KeyKOS [Bromberger et al., 1992] and EROS [Shapiro et al., 1999]. Besides its assurance story, seL4’s most characteristic aspect is its isolation-oriented approach to memory management, which is made policy-free by fully delegating it to user level.
Specifically, the kernel never allocates memory. After booting, seL4 hands all rights to any unused memory to the first user process in the form of capabilities to Untyped memory. The only operation supported on Untyped is to retypen into some other object type (TCB, page tables, frames etc), or to revoke of an earlier retype. That way user-level managers have full responsibility for any memory management. For example, the initial process can partition Untyped memory into several disjoint pools, and set up secondary resource managers in each partition. The partitions are then totally isolated, unless the initial process also provides access to some shared resources (e.g. frames or endpoints) to support communication.

Like any kernel operation (other than the yield() syscall which simply forfeits the remainder of the present time slice), IPC and notifications are authorised by capabilities: a thread needs an endpoint capability in order to send or receive messages, and a notification capability for signalling or collecting notifications.

Similar to other L4 kernels, the kernel not only supports basic send() and receive() operations, but also two versions of a send followed by a receive in one atomic syscall. First there is the RPC-like call(), which is typically used by a client invoking a server. When invoking call() on an endpoint, the kernel creates a single-use reply capability, which refers to a virtual, temporary reply endpoint. The kernel delivers the reply capability to the receiver listening on the endpoint, and makes the sender wait on the reply endpoint.

The second combined call is reply.receive(), which sends a message to the (implicitly supplied) reply endpoint and then makes the invoker wait on a new request on the endpoint specified in the syscall. Once used in the reply, the reply endpoint and capability are removed.

2.3.2 Scheduling

Management of time is comparatively under-developed in seL4. It presently implements the same simplistic scheduling model used in most L4 kernels for 20 years: priority-based round robin. The only controllable parameters are a thread’s priority and time slice. This is not sufficient for supporting MCS, as indicated by the examples given in the introduction.

On IPC, seL4 uses a direct process switch [Liedtke 1993] where possible, to avoid the cost of invoking the scheduler: the IPC switches context from sender to receiver, but with the receiver running on the sender’s time slice, until it replies or is preempted. When rescheduled after preemption, the server will execute on its own time slice (and after replying to the client, the latter may execute on the server’s time slice). The IPC paths which do not require scheduler invocation are implemented by separate, highly-optimised fast-path code.

This form of time-slice donation [Steinberg et al. 2010] has been criticised as inappropriate for RT systems [Ruocco 2008], as time is not accounted properly. Consequently, the Fiasco L4 kernel allows the sender to specify whether donation is permitted.

However, even without time-slice donation, traditional L4 scheduling is problematic. Consider a typical scenario of two clients, A, B, invoking server S. Both clients have the same priority, which is lower than the server’s, and the same time slice length, so they ought to get equal amounts of time. Assume client A requests long-running operations from S, while B’s requests are short. The server’s time is not accounted against the clients, and A gets a much higher share of the system than B. Furthermore, if the scheduler is invoked on each IPC, A and B will alternate execution after each server invocation, making it very difficult to reason about the progress of individual tasks. Alternatively, if A continues executing after the invocation of S returns, then A can effectively deny B’s service by invoking S in a tight loop.

In summary, the L4 model of managing time is unsatisfactory no matter how it is implemented. The bandwidth-inheritance approach taken in some kernels [Steinberg et al. 2010] is not a good solution either for the reasons explained in Section 2.2 on the one hand there is the general issue of complexity and poor priority-inversion bound of inheritance. On the other hand, inheritance offers no policy flexibility on managing overruns in servers. Additionally, while Fiasco’s implementation of bandwidth inheritance allows for bounded priority inversion, it violates temporal isolation: A is allowed to consume B’s budget.

2.4 Summary

We want a model that is simple enough to be suitable for seL4, provides temporal isolation, and provides freedom in the implementation of policies for dealing with isolation violations. At the same time, it must continue to support all existing or anticipated use cases of the kernel.

3 Scheduling Model

We now present a scheduling model for seL4 which satisfies all requirements for MCS stated in Section 1. It is based on a small number of abstractions, namely

- periodic threads with hard CPU bandwidth limits
- scheduling contexts
- timeout exceptions
- notion of criticality in addition to priority & explicit mode switches.
Our model matches the approach known as static mixed criticality in scheduling theory [Baruah et al., 2011], which provides appropriate tools for analysis.

### 3.1 Execution-time limits: Budgets

A key observation from [Section 2.1] is that pure priority-based scheduling cannot satisfy the requirements of MCS, and we need a mechanism for temporal isolation. To achieve this we introduce the notion of a budget, which is a hard limit on the time a thread can consume during a period. The ratio of budget over period is the limit of CPU bandwidth a thread can consume.

Budgets are similar to the reservations introduced in [Section 2.1] except that the kernel makes no guarantee that any bandwidth is achieved, only that the limit is not exceeded. This makes admission control a user-level responsibility, avoiding any policy in the kernel about whether admission should be determined on- or off-line, should be static or dynamic, or should be hierarchical of flattened [Lackorzynski et al., 2012]. In particular, the system designer may decide to trust a particular task not to use its budget (except in emergencies) and perform the schedulability analysis based on that knowledge.

Despite providing weaker guarantees, budgets are a more powerful concept than reservations. Specifically, if a set of reservations is schedulable, i.e. admission control succeeds, then budgets will produce the same schedule, i.e. they behave like reservations. If, however, the total is not schedulable, but the sum of all budgets above some threshold priority \( p \) is, then all budgets of tasks whose priority exceeds \( p \) still behave like reservations, but nothing of priority \( \leq p \) is guaranteed any CPU time.

This property allows us to safely overload a system with predictable outcomes and without the kernel performing any admission control. We will see later how this example of less is more allows us to support MCS.

Specifically, we replace the kernel’s notion of a time slice by two new attributes: period and budget. The budget is less than or equal to the period and the ratio specifies the maximum share (utilisation) of the CPU the thread can possibly get. This is essentially the model of sporadic servers introduced by Sprunt et al. [1989], except that we use budgets instead of reservations.

The operation of the seL4 scheduler changes only slightly: it still picks the highest-priority runnable thread, using round-robin within a priority. The difference is that when the kernel schedules a thread, it sets a timer to enforce the budget, and a thread whose budget is expired is no longer runnable. The period specifies when the thread’s budget is replenished, thus making it runnable again. Figure 3 shows some examples.

Similarly to the budget not guaranteeing any time, the period does not guarantee that a thread is actually scheduled periodically (which depends on the priorities, periods, and budgets of all other threads with the same or higher priority). Note also that this model exactly emulates the existing seL4 scheduler when all budgets are full: if every thread has a budget that is equal to its period, the period has the same semantics as the time slice used to have.

### 3.2 Scheduling contexts

In order to provide better control over the time resource, we introduce a scheduling context (SC) object that grants access to time. Instead of a time slice, a thread (in its TCB) holds a scheduling context capability (scCap), without such a valid scCap, the thread is not runnable.

The SC consists of the period, budget pair introduced above and thus represents the maximum bandwidth a thread may consume. The semantics of SCs are equivalent to hard reservations in Linux/RK [Rajkumar et al., 1998], in that once the budget is exhausted, no thread can run on that SC until it is replenished, however they differ in two ways. First, we only allow one thread per SC at a time, but SCs can be passed between threads via IPC for cooperative scheduling. This allows for a minimal, single level scheduler in the kernel. Second, the kernel does not conduct an admission test. Our SCs differ from those of NOVA [Steinberg et al., 2010] in that priority remains a thread attribute instead of being associated with an SC, and we allow only one SC per thread.

SCs are like other seL4 objects, in that any thread that can allocate memory can create them. However, setting the budget requires special privilege, as creating budgets amounts to control over the right to consume CPU time. It must be authorised by a capability.

We use an approach that is analogous to managing interrupt sources in seL4. Specifically, there is a per-core virtual scheduling-control object, represented by the sched_control capability. This capability must be presented when setting the budget of an SC. The kernel creates this capability at boot time and hands it to the initial task as part of the startup protocol.

SCs provide a clean solution to the shared-server ac-
Figure 3: Examples of thread schedules. P=priority, T=period, B=budget, U=max. utilisation, u=actual utilisation.

(a) Two periodic RT threads plus one best-effort thread running in slack time.

(b) Three full-budget threads scheduled as in traditional L4.

table

| P | T | B | U | u | Schedule |
|---|---|---|---|---|---------|
| 3 | 5 | 1 | 0.2 | 0.2 | ...     |
| 2 | 10 | 5 | 0.5 | 0.5 | ... |
| 1 | 20 | 20 | 1.0 | 0.3 | ... glucose | 2 | 1 | 1.0 | 0.5 | ... |

\[(a) Two periodic RT threads plus one best-effort thread running in slack time. \]

\[(b) Three full-budget threads scheduled as in traditional L4. \]

counting dilemma outlined in Section 2.3.2. We allow a (server) thread without an scCap to wait on an endpoint, we call this a **passive server**. If a client sends a message to this endpoint, the IPC will transfer the client’s scheduling context to the server, which then executes on the client’s *borrowed* budget. The SC returns to the client when the server replies to the client request. An example is given in Figure 4, where an SC-less resource server has two clients, each holding an SC (indicated by the clock dial representing a CPU bandwidth bound). For security, the sender must agree to the donation, the IPC will fail if the receiver has no SC but the sender is unwilling to lend its own. No SC transfer takes place if the receiver has its own SC (active server).

A passive server can trivially implement the immediate priority ceiling protocol introduced in Section 2.2, by setting its priority to the ceiling of priorities of all clients. As any client needs a send capability on the server’s endpoint, usermode managers can control access to the server, and thus enforce the priority ceiling. We discuss in Section 3.4 how we deal with the server running out of budget.

```c
notification_t ntfn;
sched_context s_sc;
tcb_t s_tcb;
void init() {
    // bind SC to TCB
    bind(s_sc, s_tcb);
    // create server thread
    start_thread(s_tcb);
    // block and allow
    // server to run
    wait(ntfn);
    // server initialised
    // convert to passive
    unbind(s_sc);
}
```

Figure 5: Passive server initialisation.

Passive servers must be initialised with an initialisation SC and then communicate to the initial task when they are done, such that the SC can be removed. We support this with a new system call `signal_receive()`, which combines signalling a notification with an IPC receive. Figure 5 shows how initialisation works in principle.

Call-reply&wait IPC with SC transfer avoids invoking the scheduler or updating accounting data during IPC, and thus retains the low overhead of the direct process switch optimisation. It has in fact many of the properties of a migrating thread model [Ford and Leprau 1994], specifically it avoids having multiple schedulable entities for what is logically a single-threaded operation. The advantage over migrating threads is that the kernel does not have to provide stacks on the fly, and thus is free of policy decisions such as determining stack sizes, charging for memory, whether to cache stacks. Instead, our model requires explicit user-level management of stacks through thread objects.

### 3.3 Managing thread execution

A periodic thread needs to suspend itself when it has finished processing for the current period. It does so by calling `yield()` on its own SC. An event-triggered (sporadic) thread instead waits on its IRQ notification. Of course, even if the notification is signalled (by an IRQ or another thread), the sporadic thread will only execute if it has budget.

Sometimes explicit changes of a thread’s priority are needed, e.g. when implementing IPCP without encapsulating the critical section into a separate server. In order to change a thread B’s priority, thread A must hold a capability to B’s TCB. In order to prevent arbitrary priority changes, we re-introduce the concept of a *maximum controlled priority* (MCP) that was used in early L4 versions [Liedtke 1996]. Specifically, A cannot *raise* any thread’s priority, including its own, higher than A’s MCP. Note that this does not stop A from having a priority higher than its MCP, but some other thread must have set it up.

We add two further operations on scCaps in order to allow fine-tuning scheduling decisions. The first, `consume()`, obtains the total time accounted the designated scheduling context since the last such enquiry.

3Authorising `yield()` with an SC capability removes seL4’s previous anomaly of having a syscall that requires no capability to execute. `yield()` can also be called on another thread’s SC, cancelling that thread’s current budget. However, we do not claim that there is a good use case for this.
The second, `yieldto()`, allows user level to manipulate the kernel’s scheduling queues. When invoked on a scCap, and the designated SC is presently associated with a thread whose priority does not exceed the callers MCP, and the thread has budget available in its present period, then that thread is moved to the head of the ready queue of its priority. This ensures that it is the next thread to be scheduled if no higher-priority threads are runnable. Invoking `yieldto()` implicitly invokes `consume`, i.e. it returns and resets the time accumulated on the SC.

### 3.4 Budget overrun

We provide **timeout exceptions** in order to detect budget overrun, analogous to seL4’s treatment of other exceptions. An seL4 thread already has an exception endpoint. If an exception is triggered, the kernel sends a message to the appropriate endpoint on thefaulting thread’s behalf. An exception handler waiting on the endpoint will then receive the message and take appropriate action. By replying to the exception message, it unblocks the faulting thread (possibly after adjusting its instruction pointer to skip an emulated instruction). In practice, many threads share the same exception endpoint (and thus handler).

We extend this model by adding a timeout-exception endpoint: the kernel sends a message to that endpoint when the thread exceeds its budget, and the handler can take appropriate action, which may include adjusting the faulting thread’s budget. If the handler increases the budget and then replies to the fault message, the thread will continue to run on the remainder of the enlarged budget. A thread without a timeout-exception endpoint is simply rate limited.

Timeout exceptions allow recovering from priority/criticality inversions, as possible in the passive resource server of Figure 4. If the server’s borrowed SC runs out of budget, its timeout handler can implement appropriate policy, such as letting the server complete the request on an emergency budget, forcing a reset or roll-back, possibly coupled with taking some additional safety precautions prejudicial to C1, such as suspending C1 or affecting a criticality mode switch. This is in contrast to kernel-implemented helping schemes, which implement a specific policy.

### 3.5 Criticality mode switches

Another case of budget overrun is the system shown in Table 1 consisting of three **HIGH** tasks (pink) and two **LOW** tasks (blue), plus T0 which runs in slack time. With T4’s **LOW** budget of 2 units, this system is RMPA schedulable — the RMPA utilisation bound for 5 tasks is 74% — so all tasks will meet their deadlines.

| C | P | T | B | U  |
|---|---|---|---|----|
| T3 | 1 | 6 | 10 | 2  | 0.20 |
| T4 | 1 | 5 | 20 | 2/7| 0.10| 0.35 |
| T1 | 0 | 4 | 25 | 5  | 0.20 |
| T2 | 1 | 3 | 40 | 4  | 0.20 |
| T1 | 0 | 2 | 60 | 6  | 0.20 |
| T0 | 0 | 1 | 100| 100| 0.00 |

Table 1: Parameters of a sample system, where T3 has a **LOW** budget of 2 and a **HIGH** budget of 7. C=criticality, P=priority, T=period, B=budget, U=utilisation.

Now assume that T3 overruns its budget, triggering a timeout exception. The handler can adjust its budget to the **HIGH** value of 7 units, however, the resulting system is no longer schedulable. Since the 4-task utilisation bound of RMPA is 75%, not only the **LOW** task T1 may miss its deadlines, but also the **HIGH** task T2.

We can repair this situation by a criticality switch that prevents **LOW** tasks, specifically T1 from competing with T4. We support this by introducing an explicit notion of a system criticality level, as well as a new thread criticality attribute. When setting the criticality system level to C, we boost the priority of all threads with criticality ≥ C by a constant amount, so that they all have priorities above any lower-criticality threads [Burns and Baruah, 2013]. In the above example, the timeout handler not only increases T4’s budget, but also raises the criticality level to one.

The lowest-priority task T0 will only run if there is slack in the system. If so, the criticality level can be reset to zero (possibly after waiting for a few of T0’s periods).

We control thread criticality changes similarly to priority changes: a thread attribute maximum controlled criticality (MCC) determines limits how a thread can change another thread’s criticality, just as the MCP limits priority changes. Setting the system criticality level requires the `sched_control` capability.

### 4 Implementation

#### 4.1 Objects and methods

We add a new 64-byte scheduling context object type, and modify global state by eight words plus the number of criticalities. In TCB objects we replace the `timeslice` by the `scCap`, add a timeout handler capability, criticality, MCP, and a number of bookkeeping fields, a total of nine extra fields. As TCB objects must be powers of two in size, this has no effect on the size of a TCB object.

We add three methods on TCBs. SCs have 5 methods, the new `sched_control` has two. There are also
three new methods for manipulating reply capabilities: the ability to set your reply slot, save another threads reply capability, and the ability to swap your reply capability with one saved earlier. This extra flexibility with reply capabilities allows for more efficient user-level scheduling via IPC, and allows the timeout handler to access the faulters reply capability, so it can unblock the client on the servers behalf. Also new is \texttt{nsbend\_wait}.

### 4.2 Scheduling algorithm

Baseline seL4 has a ready queue, which satisfies the invariant that it contains all runnable threads except the one presently executing [Blackham et al., 2012]. It is implemented as a priority-indexed array of queues. A two-level bitfield of occupied priorities ensures O(1) access.

The main change required to the existing seL4 scheduler is the addition of a release queue. A thread whose budget expired before its period is up is removed from the ready queue and inserted into the release queue. This retains the existing invariant for the ready queue, while the release queue is characterised as holding all threads that would be runnable but are presently lacking budget. The queue is ordered by the time of the threads next budget refresh, i.e. the time their next period is up.

Whenever the kernel schedules a thread, it sets the timer to fire when the threads SCs remaining budget is due to expire, or for the next wake-up time for the head of the release priority queue (whichever is first). If an SC switch occurs, because either fires or the thread blocks without an SC transfer, the consumed time is subtracted from the SCs budget and added to the accumulated time.

On kernel entry (except on the fastpath, which never leads to an SC change or scheduler invocation) the kernel updates the current timestamp and stores the time since the last entry. It then checks whether the thread has sufficient budget to complete the kernel operation. If not, the kernel pretends the timer has already fired, resets the budget and adds the thread to the release queue.

This adds a new invariant that any thread in the scheduling queues must have enough budget to exit the kernel. This makes the scheduler precision equal to the kernel’s WCET, which for seL4 is known (unlike any other protected-mode OS we are aware of).

Threads are only charged if the scheduling context changes, in order to avoid reprogramming the timer which can be expensive on many platforms. Else, the timestamp update is rolled back by subtracting the stored consumed value from the timestamp.

### 4.3 Criticality

A core integrity requirement of MCS is that the timeliness of HIGH tasks is unaffected by low tasks. This includes the mode switch: its cost must not depend on the number of LOW tasks in the system. We implement criticality as follows.

The kernel supports base priorities in the range $[0, 2^{2p} - 1]$, where $N_p$ is a kernel build option. The base priority is the threads actual priority at system criticality level zero. The number of criticality levels, $N_{crit}$, is also a build option. Typically, it is a small number, e.g. [\texttt{RTCA}] specifies five levels. We require that $N_{crit} \times 2^{2p} \leq 1024$.

For each criticality level the kernel maintains a queue of threads, threads that are explicitly suspended (as opposed to out of budget or blocked in IPC) are not in any criticality queue.

When system criticality changes from $C$ to $C'$, the kernel iterates through the criticality queues from $C'$ to $N_{crit} - 1$. For each thread in those queues, the kernel changes the present priority $P$ to $P_0 \land (C' \ll 8)$. This ensures that the priority of all HIGH threads is above those of all LOW ones (with respect to $C'$).

The per-priority ready queues are doubly-linked lists of TCBs, so moving a thread from one queue to another is a constant-time operation. Hence, the total time for the priority adjustments is proportional to the number of threads at criticality $C'$ or higher.

If during a criticality increase the kernel detects any threads that are running on a borrowed scheduling context (comparing tcb->sc->home to tcb), and the SCs owner is LOW (tcb->sc->home->crit $\leq C'$), it generates timeout exception for that thread. This allows a server to abort any operation on behalf of a LOW thread. If the thread running on an SC borrowed from a LOW thread has no timeout handler, it will complete normally. In this case, the worst-case blocking time is the worst-case server request time, plus the cost of the mode switch.

|               | Co-operative | Preemptive |
|---------------|--------------|------------|
| Shared SC     | IPC          | Timer notifications |
| SC per TCB    | Signals      | Timeout exceptions |

Table 2: Mechanisms for user-level scheduling

### 4.4 User-level scheduling

The kernel provides fixed-priority scheduling with budgets. This is a particular (although quite flexible) policy. Fortunately, our mechanisms allow us to implement very general policies, as indicated in Table 2.

For example, cooperative scheduling with an arbitrary policy can be implemented with a shared SC, where the threads cooperate via IPC, or per-thread SCs, where synchronisation is via notifications (although it is unclear
why one would want the latter). Pseudocode for both variants is shown in Figure 6.

Similarly, arbitrary preemptive scheduling policies can be implemented. Figure 7 shows pseudocode for schedulers with shared or per-thread SCs. The shared-SC case uses one SC for all threads, and a separate one for the timer.

```c
void coop_sched_s() {
    reply_recv(ep);
    p = t;
    t = pick_thread(p);
    swap_caller(t, p);
}

void coop_yield_s() {
    // yield
    call(ep);
}

void coop_sched_m() {
    t = pick_thread(t);
    signal(t->ntfn);
    yieldTo(t->sc);
}

void coop_yield_m() {
    // yield
    wait(ntfn);
}
```

Figure 6: User-level cooperative scheduler and thread yield function using a shared SC (left) and per-thread SCs (right).

```c
void pr_schd_s(prev) {
    // wait for timer
    wait(timer);
    t = pick_thread();
    // change sc over
    swap_sc(t, prev);
    program_timer();
    ack_irq();
}

void pr_schd_m() {
    // wait for timeout
    recv(ep);
    t = pick_thread();
    // place at head
    // of prio queue
    yield_to(t);
}
```

Figure 7: User-level preemptive scheduler with shared (left) and per-thread (right) SCs.

5 Evaluation

We conducted our evaluation on two machines, both configured to use one core:

- **Sabre**: 1 GHz ARM Cortex A9 system on chip on a Freescale i.MX6 SABRE Lite development board.
- **Haswell**: 3.1 GHz Haswell E1220v3 processor in a server machine running in 32-bit mode (64-bit seL4 is in development).

5.1 Microbenchmarks

5.1.1 Kernel microbenchmarks

Figure 8 shows the cost of the (performance-wise) most important kernel operations of our present implementation compare to the baseline seL4 kernel. Latency of the main IPC send+receive operations increases by three cycles (call) and by 12–20 cycles (reply&wait). These are the result of extra checks on the fastpath to accommodate scheduling contexts and ordering IPC, but the increase in cost is clearly negligible. The same can be said for signalling a notification.

The actual cost of the model can be seen in the IRQ and scheduler latency. Part of that is due to the need to reprogram the timer to enforce the budget, which is needed on every scheduler invocation, but also on an IRQ, as this normally unblocks a waiting handler. We measure the cost of reprogramming the timer to be 55 cycles on the Sabre, but about 200 cycles on the Haswell.

The rest of the increase is the result of the significant extra code from dealing with scheduling contexts. Note that scheduling is considered an expensive operation in seL4, and happens much less frequently than IPC.

5.1.2 Mode switch

```plaintext
| Criticaiity | Threads boosted | ARM up | ARM down | x86 up | x86 down |
|-------------|----------------|--------|----------|--------|----------|
| 3           | 4              | 1.4μs  | 1.7μs    | 0.4μs  | 0.5μs    |
| 2           | 12             | 2.4μs  | 2.4μs    | 0.5μs  | 0.6μs    |
| 1           | 28             | 4.3μs  | 3.7μs    | 0.8μs  | 0.7μs    |
```

Table 3: Results of switching from criticality level 0 to the criticality listed in column 1. Column 2 shows the number of tasks that need boosting. Standard deviations are no more than 2%.

To evaluate the cost of changing the system criticality level, configure the kernel with 256 priorities and 4 criticality levels (0–3). We then set up a system with 60 threads, of which 32, 16, 8 and 4 have criticality 0, 1, 2 and 3 respectively.

Table 3 shows the cost of switching criticality level between zero and one of the other levels. For each data
Table 3 shows the cost of switching criticality level between zero and one of the other levels. As the table shows, when switching to level three, the three threads at that level need to be boosted, while a switch to level one requires boosting all 28 threads of criticality greater than zero.

The results show that a mode switch is fairly fast, around 1,500 cycles on both platforms as long as the affected number of threads is small (which is to be assumed for \texttt{HIGH} threads), and cost is roughly linear in the number of threads to be boosted. This is important, as the schedulability analysis must allow for that cost. However, the numbers shown in Table 3 are hot-cache (best-case) numbers, while the criticality analysis must be based on WCET.

Table 3 shows the number of threads for each criticality and the results of the microbenchmark.

5.2 Case studies

5.2.1 Linux CFS

As an example of a complex dynamic-priority scheduling policy implemented at user level, we implement a version of Linux’ so-called \textit{completely fair scheduler} (CFS). The implementation uses a red-black tree and calls to \texttt{consumed()} to adjust the weights. The scheduler runs one \texttt{sel4} priority above its clients.

Figure 9 shows the scheduling cost for two scenarios, shared and per-client SDC. Cost is measured by taking a time stamp in the client, which then calls \texttt{yield()}, with another time stamp taken right after (in the next client thread).

We also show the cost of the same operations under Linux, which takes about 50–60% of the time. However, this turns out to be mostly the syscall cost, as the Linux \texttt{yield()} bypasses the scheduler. So, our user-level implementation looks quite competitive.

5.2.2 EDF scheduler

As a second scheduling policy we implement EDF at user level, this time only the scenario with a shared SC. The results are shown in Figure 10. The standard deviations are very big, especially on the Haswell platform. This is not unexpected, as the amount of work EDF has to do on each scheduling operation is very sensitive to the present state of the deadline and release queues. The scheduler may have to release threads, reprogram the timer for the next release, ack the previous interrupt and IPC the next thread, or resume a preempted thread with \texttt{yieldto}.

We used the \texttt{randfixedsum} \cite{emberson2010} algorithm to generate 10 EDF task sets for each of the 10 data points, with periods between 10–100ms. Each task set ran 1000 times, for 10,000 runs for each data point. A better metric in this case is the minimum scheduler time, shown in the figure as “Sabre min”, “Haswell min”. It is reasonably stable around 2 \mu s for the Sabre, and 0.5–0.9 \mu s for the Haswell platform. This is an excellent result: Cerqueria and Brandenburg \cite{cerqueria2013} measured the latencies of various in-kernel Linux schedulers on a Xeon X7550 platform and found the minimum to be around 1.5 \mu s for all schedulers. While comparisons across different hardware must be taken with a grain of salt, the fact that latencies of our user-level implementation is a factor four less indicates that our performance is competitive.

5.2.3 Network server

In order to demonstrate temporal isolation, we use a network benchmark, specifically the Yahoo! Cloud Serving Benchmarks (YCSB) \cite{cooper2010}. We run this against a server using the Redis key-value store \cite{redis}.

The server setup is shown in Figure 11. Dashed arrows show synchronisation operations through notifications (semaphores) indicated by flags, with coloured, broken single arrows indicating the direction of the signal. The OS server, which contains the IP stack and is
implemented as a passive server, presents a POSIX interface, which is implemented by an RPC protocol through an endpoint. The (active) Redis server invokes the OS server (coloured, solid single arrow), which then runs on Redis’ scheduling context. Redis and the OS share a buffer for passing bulk data (black, solid double arrows). The OS also shares a buffer with the Ethernet driver, which uses a second notification (red) for signalling completion to the OS. That notification is “bound”, meaning the signals are delivered to the waiting OS as an IPC apparently coming from the endpoint.

| Thread | Prio | Period | Budget |
|--------|------|--------|--------|
| Hog    | 254  | 1 ms   | variable |
| Driver | 253  | 2 ms   | 2 ms   |
| Redis  | 252  | 1 s    | 1 s    |
| OS     | 252  | - -    | - -    |

Table 4: Scheduling parameters of network server setup.

Not shown is a separate CPU hog thread, which does not communicate with this setup, but is competing for CPU time. The hog runs at highest priority (254) with a 1 ms period. The Ethernet driver runs at priority 253.

We use the budget of the hog to control the amount of time left over for the server configuration. Figure 12 shows the bandwidth achieved by the YCSB-A workload as a function of the available CPU bandwidth (i.e. the complement of the bandwidth granted to the hog thread). The figure also shows the total CPU idle time.

The graph shows that the server is CPU limited (very low idle time) and consequently throughput scales linearly with available CPU bandwidth.

5.2.4 Server rollback

As an example of a shared server running out of budget, we implement the scenario of Figure 4 of a passive server with two clients. The server is providing an encryption service using AES-256 using a block size of 16 bytes. The server alternates between two buffers, of which one always contains consistent state, the other is dirty during processing.

When the server runs out of budget, its timeout fault handler gets invoked. It rolls the server back to the last consistent state and makes it ready for the next client.

We measure rollback time, from the time the fault handler is invoked, until the server is ready for the next request. Given the small amount of rollback state, this measures the baseline overhead, for servers with more state, the handling that state would have to be added.

We run this on the Sabre and find a mean rollback time of 12 μs, with a 31% standard deviation on 12 runs with a cold cache. The individual times fluctuated between 9 and 24 μs.

6 Conclusions and Future Work

Mixed criticality systems are gaining traction in avionics and the automotive sector, due to the SWaP issues created by mushrooming functionality. In order to get the full benefit of MCS, we need an OS supporting strong, but asymmetric temporal isolation. Inherent in the notion of MCS is also a requirement for high assurance.

While there is a wealth of theory about MCS, little of it is implemented in more than a proof-of-concept, certainly not in a high-assurance OS. We have identified a model for temporal resource management that lends itself to efficient and policy-free implementation in a high-performance and high-assurance OS. We have implemented this in seL4, and have demonstrated that the base model supports the efficient implementation of a range of different scheduling policies, and allows efficient handling of various emergencies.

Presently the main limitation of the work is the restriction to a single core, which is the target of future work, as is the formal verification of the real-time seL4 kernel.
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