Can We Prove Time Protection?

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ABSTRACT

Timing channels are a significant and growing security threat in computer systems, with no established solution. We have recently argued that the OS must provide *time protection*, in analogy to the established memory protection, to protect applications from information leakage through timing channels. Based on a recently-proposed implementation of time protection in the seL4 microkernel, we investigate how such an implementation could be formally proved to prevent timing channels. We postulate that this should be possible by reasoning about a highly abstracted representation of the shared hardware resources that cause timing channels.

1 INTRODUCTION

Timing channels are a major threat to information security, they exist where the timing of a sequence of observable events depends on secret information [Wray 1991]. The observation might be of an externally visible event, such as the response time of a server, and might be exploitable over intercontinental distances [Cock et al. 2014]. Or it might only be locally observable, i.e. by a process or VM co-located on the same physical machine, which still enables remote attacks, if the observing process has access to the network and is controlled by a remote agent. The seriousness of the threat was recently highlighted by the Spectre attacks [Kocher et al. 2019], where speculatively executed gadgets leak information via a covert timing channel.

The secret-dependence of events may have *algorithmic* causes, e.g. crypto implementations with secret-dependent code paths. Or they may result from interference resulting from competing access to limited hardware resources, such as caches; there exists a wide variety of such *micro-architectural channels* [Ge et al. 2018b].

Whether algorithmic or micro-architectural, those channels represent information flow across protection boundaries, i.e. *the boundaries are leaky*. Ensuring the security of these boundaries should be the job of the operating system (OS); however, no contemporary, general-purpose OS seems to be capable of if. Clearly, this is not an acceptable situation, and we have recently called for OSes to provide *time protection* [Ge et al. 2018a] as the temporal equivalent of the well-established concept of memory protection.

Memory protection is a solved problem: the formal verification of seL4 proved, among others, that the kernel is able to enforce spatial integrity, availability and confidentiality [Klein et al. 2014]. This categorically rules out information leakage via *storage channels* (provided that the kernel is aware of the state that can be used for such channels). However, the approach taken in the seL4 verification has no concept of time, and therefore cannot make any claims about *timing channels*.

Our aim is to rule out timing-channel leakage just as categorically as information flow via storage. Put differently, *we aim to formally prove correct implementation of time protection*. This paper investigates the feasibility of, and prerequisites for, achieving the stated aim. Obviously we would not bother writing this paper if we were not convinced that it is feasible to achieve our aim, under certain conditions, which come down to hardware satisfying certain requirements. We have recently demonstrated that not all recent processors satisfy these requirements, resulting in a call for a new, security-oriented hardware-software contract [Ge et al. 2018a]. We claim that, for hardware that honours this contract, we will be able to achieve our aim of proving time protection, and thus eliminate micro-architectural timing channels.

Note that other physical channels, such as power draw, temperature, or acoustic or electromagnetic emanation, are outside the scope of this work.

2 THREAT SCENARIO

The basic problem we are concerned with is a secret held by one security domain, Hi, being leaked to another domain, Lo, which is not supposed to know it. The leaking might be intentional, by a bad actor (Trojan) inside Hi, constituting a *covert channel*. Or it can be unintentional, via a *side channel*. Note that Hi, Lo are relative to a particular secret, we do not assume a hierarchical security policy such as Bell and LaPadula [1976], and there may be other secrets for which the roles of the domains are reversed. It is the duty of the OS to prevent any unauthorised information flow, no matter what the system's specific security policy might be.

Our notion of a security domain refers to a subset of the system which is treated as an opaque unit by the system's security policy (i.e. intra-domain information flow is not restricted by the policy). In OS terms, a domain consists of one or more (cooperating) processes.

We assume that the OS provides strong, verified memory protection, and is free of storage channels, seL4 being an example. Our primary concern is microarchitectural channels, i.e. channels that exploit competition for finite hardware resources that are abstracted away by the instruction-set architecture (ISA), the classic hardware-software contract. This means that algorithmic channels are not our primary concern, but we will discuss in Sect. 4.3 how time protection can be employed to remove such channels (within limits).

Like memory protection, time protection is a black-box OS mechanism, that provides *mandatory security enforcement* without relying on application cooperation.

For realism, i.e. to ensure that contemporary hardware is at least close to satisfying the requirements of time protection (and can fully satisfy them with minor enhancements) we limit our scope in one important way: we do not (yet) attempt to prevent *covert channels* through stateless interconnects. Such channels, exploiting the finite bandwidth of interconnects through concurrent competing access, are trivial to implement: a Trojan running on one core signals by modulating its use of interconnect bandwidth, and a spy running on a different core measures the remaining bandwidth by trying to saturate the shared interconnect. Such channels can only be prevented with hardware support that is not available on any contemporary mainstream hardware.[1] We will be able to extend time protection in a fairly straightforward way, should such hardware support (or at least an accepted model for it) become available.

An obvious example of the excluded scenario would be a covert channel between two virtual machines (VMs) concurrently executing on different cores of the same processor on a public cloud. Such a covert channel is not a particular concern, as the Trojan in the victim VM does not need the co-located spy, as it can communicate by other means, e.g. modulating its network communication. Side channels are a real concern in the cloud scenario, but stateless interconnects reveal no address information. As a consequence, no such side channels have been demonstrated to date [Ge et al. 2018b], and they are likely impossible.

[1]Intel recently introduced *memory bandwidth allocation* (MBA) technology, which imposes *approximate* limits on the memory bandwidth available to a core [Intel Corporation 2016]. While this represents a step towards bandwidth partitioning, the approximate enforcement is not sufficient for preventing covert channels.

3 TIMING-CHANNEL MECHANISMS

There are two ways in which Lo may learn Hi's secret: by timing observable actions of Hi, or by Lo observing how its own execution speed is influenced by Hi's execution.

3.1 Timing own progress

This channel utilises the performance impact of interference between processes resulting from competition for shared hardware resources, especially stateful resources such as caches, TLBs, branch predictors and pre-fetcher state machines. For example, Lo's rate of progress (performance) is affected by cache misses. If Lo shares a cache with Hi (either time-sharing a core-private cache or concurrently sharing a cache with Hi's core), then the miss rate will depend on Hi's cache usage. If the cache is set-associative (which almost all of them are nowadays), then the pattern of cache misses will also reveal address information from Hi. Such address information supports the implementation of side channels with potentially high bandwidth, e.g. where the secret is used to index a table [Ge et al. 2018b].

An effective exploitation of such a channel is the *prime-and-probe* technique [Osvik et al. 2006; Percival 2005]. Here Lo fills the cache by traversing a buffer large enough to cover the cache (prime phase). After or while Hi is executing, Lo traverses the buffer again, monitoring the time taken for each access (probe phase); a long latency indicates a conflict miss with Hi's cache footprint. The address of the missing access reveals the index bits of Hi's access.

Prime-and-probe can be used as a high-bandwidth covert channel, where Hi explicitly encodes information into the memory addresses accessed, or as a side channel, where the encoding is implicit in Hi's normal execution (e.g. via a secret-derived array index). It can be used for time-shared (core-private) caches as well as caches shared between cores.

3.2 Timing Hi events

On a first glance this might seem like a silly case, why worry about a covert channel if there is an overt one, such as message passing? However, this situation is in fact common: Hi might be a *downgrader*, an entity trusted to handle secrets and decide which can be safely declassified. A common example is a crypto component, which encrypts secrets, e.g. from a web server, and publishes the encrypted text, by handing them to a network unit; this is shown in Figure 1.

In this case, the leakage might be resulting from an algorithmic channel (e.g. a crypto implementation with secret-dependent execution), a Trojan modulating the

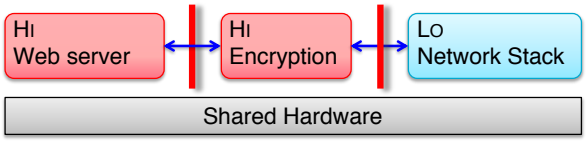

Figure 1: Encryption engine as a downgrader.

speed of the encryption (possibly via micro-architectural interference), or the server itself leaking through the timing of messages to the crypto component.

Time protection here must make execution time deterministic, meaning that message passing or context switching happen at pre-determined times. Obviously, the OS can only provide the mechanism here (deterministic switch/delivery time), not the policy (the time of the switch). This must be set by the system designer or security officer, taking into account issues like the worst-case execution time (WCET) of the encryption.

Cock et al. [2014] have proposed a possible model: a synchronous IPC channel switches to the receiver only once the sender domain has executed for a predetermined minimum amount of time. It is then left to the system designer to determine a safe time threshold.

4 DEFENCES AND LIMITATIONS

4.1 Flushing and partitioning hardware

As micro-architectural timing channels result from competition for (non-architected) hardware resources, eliminating them requires removing the competition. This means the OS must either partition those resources between security domains, or reset them to a defined, history-independent state between accesses from different domains.

Resetting, e.g. flushing caches, only helps where accesses from different domains are separated in time, i.e. for time-shared resources. In other words, resetting only works for resources that are private to an execution stream. In the absence of hyperthreading, this applies to core-local resources, such as the L1 caches, private L2 caches (on Intel hardware), TLBs, branch predictors, and core-local prefetchers.

Partitioning is the only option where concurrent accesses happen, i.e. for caches shared between cores. It would also be the only option for core-local state when hyperthreading is enabled. However, no mainstream hardware supports partitioning of hardware resources between hyperthreads, and such partitioning would seem fundamentally at odds with the concept of hyperthreading, which is based on improving hardware utilisation by sharing. Consequently there are a plethora of side-channel attacks between hyperthreads [Ge et al. 2018b]. We have to conclude that *hyperthreading is fundamentally insecure*, and multiple hardware threads must never be allocated to different security domains (multi-threading a single domain is not a security issue).

Partitioning of shared (physically-addressed) caches is possible without extra hardware support by using page colouring [Kessler and Hill 1992; Liedtke et al. 1997; Lynch et al. 1992]. This uses the fact that the associative lookup of a large cache forces a page into a specific subset of the cache, so only pages mapping to the same subset, said to have the same colour, can compete for cache space. By ensuring that different security domains are allocated physical page frames of disjoint colours, the OS can partition the cache between domains. Modern last-level caches have at least 64 different colours.

In general, *micro-architectural timing channels can be prevented if all shared hardware can be either partitioned of flushed by the OS*, with flushing the only option where accesses are concurrent. Together with a few other conditions outlined by Ge et al. [2018a], these form part of a security-oriented hardware-software contract, called the aISA (augmented ISA), that allows the OS to prevent timing channels. The ISA alone is an insufficient contract for ensuring security [Heiser 2018; Hill 2018].

4.2 Implementing time protection

We have recently proposed an implementation of time protection in seL4, for hardware that conforms to a security-oriented aISA [Ge et al. 2019]. It uses cache colouring to partition shared caches. As even read-only sharing of code is sufficient for creating a channel [Gullasch et al. 2011; Yarom and Falkner 2014], we also colour the kernel image. This is achieved by a policy-free *kernel clone* mechanism, which allows setting up a domain-private kernel image in coloured memory.

We flush time-shared micro-architectural state on each domain switch (but not on intra-domain context-switches). For writable micro-architectural state (e.g. the L1 data cache), the latency of the flush is itself dependent on execution history (number of dirty lines), which would create a channel. We avoid this channel by padding the domain-switch latency to a fixed value. For generality (see Sect. 4.3) we make determining the padding time not the job of the OS, but an attribute of the switched-from security domain, controlled by the system designer. Specifically, we specify that the next domain will not start executing earlier than the previous domain's time slice plus the padding time.

The padding time should obviously be at least the worst-case latency of the flush, but also needs to account

for any delay of the handling of the preemption-timer interrupt by other kernel entries (resulting from system calls or interrupts).

Finally, interrupts could also be used as a channel, if the Trojan triggers an I/O such that its completion interrupt fires during Lo's execution [Ge et al. 2019]. We prevent this by partitioning interrupts (other than the preemption timer) between domains, and keep all interrupts masked that are not associated with the presently-executing domain.

4.3 Preventing algorithmic channels

Padding is a general mechanism that can also be used to prevent algorithmic channels. In the scenario of Figure 1, we can pad the execution time of the downgrader to a safe value (an upper bound of its execution time). In practice, this is very wastive if padding is done by busy looping. To make it practical, another Hi process should be scheduled for padding. Obviously, that interim process must be preempted early enough to allow the kernel to switch domains without *exceeding* the pad time (as this might introduce new channels). This may not be straightforward to implement, but it is clearly possible.

5 PROVING TIME PROTECTION

At first glance, one might expect that proving time protection is a hopeless exercise. After all, the precise interaction between microarchitectural state and execution latency is unspecified for modern hardware platforms, and the latency of some instructions may vary by orders of magnitude depending on hardware state. Formally reasoning about precise execution latencies is therefore infeasible [Klein et al. 2011].

However, we argue that reasoning about the exact latency of executions is unnecessary. **The key insight is that these channels are effected by shared hardware resources, and if we can prove that no sharing happens, there can be no timing channels.** Consequently, proving temporal isolation requires formal models of microarchitectural state, but these can be kept abstract, providing only detail to identify resources that need to be partitioned (and how such partitioning is performed), and state that must be reset (and how to reset it). That is, we do not need to know how long an instruction will take to execute, only which micro-architectural state its execution time depends on and how this state behaves wrt. partitioning and flushing.

For partitionable state, temporal isolation becomes a *functional* property (namely an invariant about correct partitioning) that can be verified without any reference to time, meaning existing verification techniques apply.

For state that requires flushing, correct application of the flush is also a functional property. As mentioned in Sect. 4.2, the latency of flushing operations themselves needs to be hidden by the OS, by padding its execution. Correct padding can be verified with a relatively simple formalisation of hardware clocks, which allows verifying padding time by simply comparing time stamps, reducing this to a functional property as well.

Once timing-channel reasoning is reduced to the verification of functional properties, it should be possible to integrate it into existing proof frameworks of storage-channel freedom, such as seL4's information flow proofs [Murray et al. 2013].

Indeed, under this approach **timing-channel reasoning is transmuted into reasoning about storage channels**, reducing it to a solved problem, and also enabling reasoning about timing-channels without reference to precise execution time. This possibility may seem surprising, but it is known that the distinction between storage and timing channels is not fundamental, but refers to the mechanisms used for exploitation [Wray 1991]. In our case we transform the temporal interference problem into a spatial one, by reasoning about the shared hardware resources which the channels exploit.

5.1 Hardware formalisation

Carrying out these proofs requires a model of the shared hardware resources (the *microarchitectural model*) that influence execution latencies, as well as a simple model of a hardware clock (the *time model*) to allow reasoning about elapsed time intervals. Naturally these models are interrelated: how much an execution step advances the hardware clock naturally depends on the microarchitectural state that influences execution time.

Crucially, a precise description of this interaction is not necessary. Instead, the interaction can be faithfully yet feasibly modelled as follows. Firstly, the microarchitectural model must delineate the partitionable state from the flushable state, and all microarchitectural state must be partitionable or flushable (Sect. 4.1). Secondly, the time model, which captures how far time advances on each execution step, is defined as a *deterministic yet unspecified* function of the microarchitectural state. Then, when the microarchitectural state is properly partitioned and flushed, one can prove that a security domain's execution time cannot be influenced by other domains (see Sect. 5.2 below).

This construction neatly reflects the basic assumptions that (i) the hardware provides sufficient mechanisms to partition and flush microarchitectural state between security domains, that (ii) such mechanisms

work correctly, and that (iii) these account for all microarchitectural state that influences execution time.

5.2 Information-flow proofs

With these models in hand, time protection can then be proved by showing that there is no way in which the execution of one domain can affect the execution timing of another domain.

Specifically the proofs must show that all resource partitioning and flushing is applied at all times and not bypassable, and that domain-switches (flushing) is correctly padded to a constant amount of time (under the assumption that the padding value, obtained by a separate analysis, is sufficient). These proofs can then be integrated with existing storage-channel freedom proofs to derive the absence of timing channels as follows.

Without loss of generality, fix some domain (Lo) and consider one of its execution steps for which we show its timing cannot be influenced by another domain (Hi). There are two possibilities: (Case 1) either it is an ordinary user-mode instruction, or (Case 2) it is a trap (a system call, exception, or interrupt arrival). For Case 1, the execution time given by the time model will be affected by the shared hardware resources in the microarchitectural model. Recall that this effect can be modelled by an unspecified deterministic function from the state of the microarchitectural model to an elapsed (symbolic) time value. For an individual instruction this function will examine the state of the instruction cache, namely the cache set identified by the program counter, and the state of the data cache for any memory address accessed by that instruction. Since the access does not fault (otherwise it would be a trap), all such memory accesses must lie within the physical memory of the current domain and thus within areas of the cache that cannot be affected by other partitions (due to correct cache partitioning by the kernel, or correct flushing, e.g. for the on-core L1 cache).

A similar argument applies to other microarchitectural resources. Thus the resulting execution time cannot be affected by other partitions.

For Case 2, we distinguish two sub-cases: The trap is either (Case 2a) a system call or exception, or (Case 2b) it is the arrival of a timer interrupt signalling a switch to the next domain. For Case 2a, the execution time depends on the state of the instruction cache wrt. the kernel instructions executed, plus the data cache for any data accessed. However, in a partitioned system with the kernel correctly cloned as in Sect. 4.2, the former cannot be affected by other partitions and the latter accesses only data of the current domain. The only remaining state that might be accessed is global kernel data, which we will prove is accessed deterministically and whose cache state after a domain switch is independent of prior Hi activity (due to correct flushing). Thus a similar, if naturally more involved, argument applies as to the user mode case (Case 1). Incorporating general (i.e. non-timer) interrupts we believe is also possible, by partitioning the interrupt set as covered in Sect. 4.2. For Case 2b, we invoke the proof of the constant-time domain switch property. □

Note that by reflecting elapsed time as a value in the state of the time model (updated by an unspecified function of the microarchitectural model), timing-channel reasoning is reduced to storage-channel reasoning, and indeed time protection itself can be phrased and proved akin to storage-channel freedom via a suitable noninterference property [Murray et al. 2012].

5.3 TLB

The TLB is an example where the principles of partitioning and flushing can already be observed in a formal model for pure functional correctness: while not yet suitable for reasoning about timing, Syeda and Klein [2018] provide a logic for functional correctness under an ARM-style TLB. For instance, it is easy to show in this model that page tables modifications under one address space identified (ASID) do not affect TLB consistency for any other ASID. This is the kind of partitioning theorem we would make use of for timing-relevant state.

The model in this work is a high-level abstraction of the TLB proved sound with respect to a low-level model that would be infeasible to reason about directly. We propose the same for timing behaviour. Instead of reasoning about a detailed low-level architecture model with precise timing information, we only record the information needed for timing-independence.

6 CONCLUSIONS

We conclude that proving time protection should be possible with established formal methods, thanks to the key insight that they result from spatial-type microarchitectural resources, and can thus be treated as storage channels. This requires some reasoning about those hardware resources, but we expect to get away with very high-level abstractions. The key challenge is to achieve agreement on a hardware-software contract that makes it at least possible to remove timing channels. We are clearly at the mercy of processor manufacturers here!

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