The Endokernel: Fast, Secure, and Programmable Subprocess Virtualization

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Abstract
Commodity applications contain more and more combinations of interacting components (user, application, library, and system) and exhibit increasingly diverse tradeoffs between isolation, performance, and programmability. We argue that the challenge of future runtime isolation is best met by embracing the multi-principle nature of applications, rethinking process architecture for fast and extensible intra-process isolation. We present, the Endokernel, a new process model and security architecture that nests an extensible monitor into the standard process for building efficient least-authority abstractions. The Endokernel introduces a new virtual machine abstraction for representing subprocess authority, which is enforced by an efficient self-isolating monitor that maps the abstraction to system level objects (processes, threads, files, and signals). We show how the Endokernel Architecture can be used to develop specialized separation abstractions using an exokernel-like organization to provide virtual privilege rings, which we use to reorganize and secure NGINX. Our prototype, includes a new syscall monitor, the nexpoline, and explores the tradeoffs of implementing it with diverse mechanisms, including Intel® Control Enhancement Technology. Overall, we believe sub-process isolation is a must and that the Endokernel Architecture exposes an essential set of abstractions for realizing this in a simple and feasible way.

1 Introduction
The process abstraction [59] defines an interface between an application and system resources, which is heavyweight and slow for cross-component interactions. However, without process isolation, diverse multi-principle environments with large fault domains will suffer reliability and security problems [9, 18]. As a result, many attempts have been made to control access to system resources (files, address spaces, signals) through system object virtualization [20, 43, 56, 72] or kernel-level mechanisms [42, 53, 57, 65, 82]. But, all of these languages enforce rigid interfaces that, while quite expressive, are not extensible and by nature introduce threats to unrelated parts of the kernel. Beyond that, they also fail to monitor and enforce resource control at the subprocess level.

Until recently, software fault isolation [8, 44, 63, 76] was the only practical method for isolating elements within a single address space, which uses inline reference monitors. Fortunately, both hardware and software vendors have observed this trend and invested into subprocess isolation. These systems include memory and CPU virtualization that allows for multiple protection domains within a process with fast domain switching (memory protection keys [54], nested paging fast context-switch support [4, 74], etc.). The core technique is to nest a monitor into the process that efficiently enforces subprocess access control to memory and CPU state [10, 17, 24, 25, 75]. Unfortunately, these approaches only virtualize minimal parts of the CPU and neglect their domain abstraction to system resources. As a result attackers can easily break out of the sandbox. For example, suppose $s$ points at an isolated session key, then the following pseudo code would bypass any memory protection:

```python
open("/proc/self/mem","r").seek(s).read(0x10)
```

Alternatively, other intra-process isolation approaches, like Native Client [81], avoid this problem by completely eliminating access. However, this severely limits functionality of the code, and neglects a large application space. It is clear that system objects require more than all or nothing access, which says nothing about signals and IPC. Beyond these security gaps, these approaches made ad hoc guesses at the best securing abstractions, but do not provide detailed specialized protection actually useful for applications.

We believe that the core problem is the disconnect between the need for subprocess level resource control and the inability of existing solutions to extensibly and fully abstract it. The core research question of this paper is: Can we easily add an efficient and extensible security architecture that does not decrease the security of the OS while fully virtualizing access to all system resources? As such, we explore a new process model, the Endokernel Architecture\(^1\), where the monitor is nested within the process to export a lightweight virtual machine for representing and enforcing subprocess resource isolation with novel custom security languages fit for the application at hand. There are two key innovations: 1) We map the new virtual machine to the traditional process level abstractions to ensure no leaks that plagued the prior work, and 2) make it efficient and extensible to support custom protection abstractions.

\(^1\) endo-: greek for “within” and stands in contrast to exo-
We demonstrate the value of Endokernel Architecture by developing a simple yet powerful virtual ringed process architecture to address several common exploit patterns (§2.4).

Our prototype Endokernel, named Intravirt, uses memory protection keys for efficient memory isolation, with a new monitor to intercept all syscalls and virtualize signals, and is applied to implement and evaluate concurrency and multi-domain mechanisms. The monitor enforces using a new secure syscall callgate, nexpoline, which we implement using several diverse low-level mechanisms to ensure that only the monitor is permitted to invoke a syscall. To virtualize signals, Intravirt inserts a layer of indirection and maps signals to the appropriate endprocess, and has to overcome the complexity from concurrency and the subtle, poor interfaces presented by the kernel. Intravirt also identifies and prevents multi-process attacks that can bypass policies via copy-on-write address spaces. Last, while prior work indicated multi-domain separation, we identify unexplored low-level threats due to the domain switching mechanism, which we address with a secure context-switch gate.

We evaluate and compare Intravirt to a ptrace-based approach and ERIM [75] which provides only memory protection but no system isolation. We also implement and apply several exploits to test the security of our solution. Overall, overheads range from 2-40% while closing many security gaps. In summary our contributions include:

- New architecture for endprocess virtualization: Endokernel architecture that virtualizes system objects solely at user level while remaining fast and securing against new low-level attacks (§3 and §4).
- A prototype, Intravirt, that includes concurrency and scalability (§4 and §6); the nexpoline definition and exploration of several lightweight syscall monitoring mechanisms (§4.3), including a novel use of Intel® Control Enforcement Technology (CET); and a systematic representation of system object virtualization (§4.4).
- Programmable security abstractions (§5) with example Nested boxing facilities (§5.1) and novel decomposition of sudo and NGINX (§5.3), demonstrates only 5.5% overhead for securing private keys and preventing parser-bug privilege escalation (§6.2.3).

2 Motivations

Modularity is traditionally provided using virtualization, where the hardware, OS, and programming languages combine to provide the illusion that a thread of execution has the full machine. The basic abstraction for virtualization is a process, where one or multiple threads of execution share an address space and virtualized system objects. The problem is, a process is heavyweight and execution in the same address space tends to fail to apply least-authority design. As a result, malicious or compromised components will allow an attacker access to the full authority of the process.

In this section we argue that the challenges facing future application resource control are incompatible with the existing process-based isolation. While these application trends are not new [7, 19, 46, 47], innovations in hardware and software fault monitors along with nested resource virtualization open up a new opportunity for fast, safe, and extensible process virtualization. We argue that future application resource management must address the following problems: 1) lack of full virtualizability leads to incompatibility and insecurity, 2) lack of extensibility leads to hard-to-apply abstractions, 3) incomplete implementations lead to insecurity and inconclusive results, and 4) lack of hardware portability complicates integration.

2.1 Full Virtualizability

Just like the original challenge of VMMs set out by Popek and Golberg [55], current subprocess isolation suffers from incomplete virtualization. To complete subprocess isolation, we need to create a virtual machine that is a subset of the standard process while also multiplexing OS objects. Incomplete virtualization leads to direct access that breaks the isolation. Problematically, most systems that provide this type of isolation focus primarily on isolating memory but neglect control flow and data access that are made available through the system resources, including file system and address space objects. Additionally, interrupted program state and signal delivery are neglected and challenging to support.

On the other hand, some systems apply proper sandboxing, such as Native Client [81], but only provide restrictive virtual machine abstractions that fail to implement the desirable applications. For example, the sandboxed code cannot make system calls, effectively providing only all or nothing policies. Sometimes, parts of an application may need access to some but not all files, such as access to secret and sensitive data, while other components need access to other, less-sensitive data, such as configuration files. Existing interfaces for OS object virtualization are also insufficient for supporting more flexible policies at subprocess level.

2.2 Programmable Security Abstractions

Much like the Exokernel argument [20], today’s process-based isolation is inflexible. However, unlike Exokernel, the key challenge is not about exposing state for managing performance, but rather making the policy language more closely matching the needs of applications. This influences 1) Ease of use: A primary reason why fine-grained security is not applied is the complexity and diverse nature of application demands. We argue that an abstraction that works for one application won’t necessarily be the easiest to apply for another. 2) Performance. We believe that an extensible protection architecture will ameliorate these issues by putting control into application specific abstractions.

2.3 Mechanism Portability

The key problem is what are the essential elements independent of the mechanisms. It is clear that intra-process isolation mechanisms are only going to see increased exploration, which fractures the landscape of approaches for applying them. Each new system provides some properties, but how do we compare them? We believe it is necessary to establish a model that prescribes a set of clear abstractions and security properties so that diverse systems can be reasonably and
systematically applied and compared.

2.4 Mechanisms Gaps and Challenges

Several facets must be preserved to have meaningful privilege separation and compares related efforts. The key gaps and challenges are described below.

Subdomain Identifiability One solution would be to extend the kernel with subprocess abstractions. However, a userspace monitor is still necessary to track the current protection domain or else you have to transition to kernel on each switch which is prohibitively costly.

Programmability and Optimizations Having a general interfaces would be ideal but as implored by prior work (Exokernel, etc), applications tend to be severely constrained. What’s worse is that existing process abstraction needs to have a separate interface to accommodate different interaction pattern and to be efficient. Thus customizability of the abstractions is critical and most prior work don’t handle it properly.

Leaky System Objects Since OSs are unaware of subprocess domains, an untrusted portion of a application can request access and the OS will gladly service it. Although we shows several bypass attacks, the primary challenge is to systematically assesses all interfaces and to integrate them into a unified policy management interface. It is easier to reason about the policy for a relatively strict interface, but things like ioctls make it impossible to have comprehensive defenses.

System Flow Policies A basic property is that the information in a subspace should never flows in or out of system objects unless explicitly granted. However, deriving the system flows itself is hard due to system complexity. Although prior work such as Erim and Hodor shows that one can reason about the flows through a specific system object, the approach is hard to be broadened to a systematic solution.

syscall Monitor The need to monitor syscalls is clear, but how to do it is not. A deny-all policy—as used by intra-app sandboxing [49, 77, 81]—sandboxing would indiscriminately deny all access and neglect a large application space. For example, deny-all sandboxing cannot prevent Heartbleed [18]. In general, applications should be able to benefit from privilege separation while not losing functionality. Alternatively, we could modify the OS so that it recognizes and enforces endoprocesses [25, 39, 70]. Unfortunately, this introduces significant complexity as indicated by Sirius [70].

Instead of the in-kernel approach, we propose enforcing nested flow policies at the syscalls—allowing some to bypass without change, others to be denied, and the rest to be securely emulated. This is not supported by well-known systems in Linux: MBOX uses ptrace for similar protections [31], but only virtualizes the filesystem interface and is inefficient. Seccomp [13] with eBPF [21] and LSM [79] enforce syscall policies, but lack the ability to modify syscall semantics, which will require modifying the LSM hooks extensively.

Multi-Process An attacker can fork an exploited process, and access the original address space directly through load and stores instructions and access indirectly through read system calls. The endokernel must be inside the new process to ensure the protections, or the memory must be scrubbed. Prior approaches [24, 70, 75] do not consider this threat and would have to disallow fork system calls.

Signals Signals create several exploitable gaps and challenges. First, Linux exposes virtual CPU state to the signal handler including PKRU, which can be exploited by an attacker. Second, the kernel does not change the domain and will trap if not properly set up. Third, the kernel always delivers the signal to a default domain, exposing the monitor control-flow attacks. Fourth, properly virtualizing signals requires complex synchronization and modifying the semantics to be both correct, safe, and efficient. Overall, properly handling signals introduces significant complexity into the endokernel.

Multi-Threaded While existing work claims to have a design supporting multi-threading, none of them have implemented concurrency control in the runtime monitor, introducing TOCTOU attacks and memory leaks, as well as neglects to measure scalability.

Multi-Domain Prior work isolated one domain per thread but not multiple domains per thread. The challenge is that switching from the untrusted domain to the monitor exposes less data than executing an cross domain call because the stack requires tracking to ensure return integrity.

3 Endoprocess Model

The Endoprocess Model is a general purpose model for nesting a monitor, the endokernel, into the process address space, that self-isolates to create two privilege levels within the process, and presents a lightweight virtual machine, the endoprocess, to the application. The Endokernel has been designed to insert directly below application logic and directly on top of the OS and HW provided abstractions. The core methodology is to systematically identify 1) what needs to be protected, 2) how that information can be interacted with (through the CPU, memory, or OS interfaces), and 3) specify a set of abstractions that must be in place to secure endoprocess isolation. The basic goal is to identify an architecture level description that is portable and independent of the exact layers above and below to properly encapsulate the endoprocess internals. The architecture has two main elements: 1) the authority model and 2) the nested endokernel architecture that ensures isolation. Figure 2,
depicts these three elements together in the architecture.

3.1 Design Principles

We share the trusted monitor principles as outlined by Needham—tamper-proof, non-bypassable, and small enough to verify [51]—and add the following:

**Nested Separation Kernel** Address spaces and kernel interactions are slow, eliminate all OS interactions [16, 60]—i.e., pure userspace, while being smaller than a microkernel and only tolerating elements inside if they support primitive separation mechanisms with a minimal interface.

**Self-Contained and Secure Userspace** Avoid implementing system object isolation in the kernel: adding yet another security framework hacked on top of thousands of kernel objects. Nesting requires part of the mechanism to be in-process, however, certain resources could be virtualized by the OS. While that seems like it might be the best choice, if parts of the process were virtualized by the monitor and others by the OS then: 1) complexity arises in bridging the semantic gap of the abstractions, 2) bugs can arise from complex concurrency, access, and exception control, and 3) ties the endoprocess abstraction to a specific kernel implementation instead of the semantics of its interface.

**General and Extensible** The design should permit many implementations, i.e., using various hardware (MPK) or software isolation (SFI) techniques that might present valuable tradeoff points in the security-performance space. The architecture should enable safe extensibility of the security abstractions to enable custom, least-authority protection services.

3.2 Authority Model

The Endokernel represents and enforces authority based on a protection domain, called an endoprocess. As outlined by Lampson [36] and instantiated by Mondrix [78], an endoprocess must provide the basic properties of data abstraction: protected entry/return and memory isolation, while also protecting access through OS objects. Most existing work multiplexes regions of the virtual address space and protects control flow, however, these works neglect to map memory space isolation to the system objects. Thus, in addition to CPU and memory virtualization, the Endokernel also virtualizes: CPU registers, the file system, mappings, and exceptions (as implemented through signals).

**Definition 1.** An endoprocess is a lightweight virtual machine represented by (instruction, subspace, entry-point, return-point, file, mapping, process, and exception) capabilities.

Instruction capabilities specify which instructions are permitted without monitoring, and is required to fully virtualize the CPU—similar to the hosted architecture of VMMs, SFI, and Nested Kernel approaches. Explicitly representing instructions is critical as many protection models control instructions by using hardware privilege levels (rings), capability hardware, software based techniques like SFI (inline monitors), or deprivileging (static verifiers w/ runtime code integrity). The way we virtualize the CPU also influences low-level mechanisms that enforce protected entry and exits.

Memory capabilities allow an endoprocess to read, write, or execute a subspace, which is a subregion of the virtual address space. The default subspaces for each endoprocess include: stack, heap, and code. File system capabilities specify operations permitted for opening, reading, and writing runtime state through the file system. Mapping capabilities determine what changes to the address space (e.g., mmap, mprotect, etc.) an endoprocess has. Process capabilities determine what interactions are permitted between processes. Exception capabilities allow an endoprocess to securely register for and handle signals (e.g., SIGSEGV). Entry-point capabilities denote points at which an endoprocess transition is permitted, effectively converting function calls into RPCs. Return-point capabilities are dynamically generated on cross-domain calls, xcall, and require the machine to return in nested order. Each endoprocess, by default, is granted exclusive access to its own code, data, and stack subspaces. An execution context is the combination of the (endoprocess X thread context). Thus the endoprocess is similar to a traditional process by allowing multiple threads to coexist concurrently, while threads traverse endoprocesses as they execute.

**Property 1 (Endoprocess Isolation).** Each endoprocess is granted exclusive access to its code, data, and stack subspaces, guaranteed secure entry/return, mapping capabilities for its own subspaces, and capabilities to OS level interfaces unless explicitly excluded for isolating other endoprocess state.

With these capabilities, the Endokernel exposes the ability to fully virtualize each resource while restricting access to privileged in-process state (e.g., monitor memory). This is essential as many applications cannot be deployed without a certain level of access, but the monitor itself must ensure its protection.

3.3 Nested Endokernel Organization

The Endokernel Architecture is a process model where a security monitor, the endokernel, is nested within the address space with full authority. The endokernel multiplexes the process to enforce modularity in a set of endoprocesses. The first goal of the endokernel is to self-isolate, i.e., secure the endokernel state and endoprocess abstraction from untrusted domain bypass.

**In-Process Policy** The endokernel is granted full authority to all process resources, and the untrusted domain is granted access to all process resources except for the following: endokernel subspaces; memory protection (e.g., registers via NRPKIU) and direct OS call (e.g., syscalls) instructions; file system operations that would allow access to endokernel subspaces (e.g., read/write /proc/self/mem/endokernel-subspace); address space manipulation (e.g., mmap) that would expose endokernel subspaces; and signal capabilities that could otherwise use to bypass subspace isolation. Just like any kernelized system, protected gates ensure the endokernel is securely entered into when a protection domain switch occurs. This architecture is similar to and inspired by the hosted VMM architecture and Nested Kernel Architecture.

**Definition 2 (Endokernel Architecture).** An Endokernel Architecture is a split process model where the endokernel is nested within the address space.
The endokernel is responsible for exporting the basic endoprocess abstractions for all untrusted domain endoprocesses, thus enabling a new method for virtualizing subprocess resources and enforcing the following property:

**Property 2** (Complete Mediation). A non-bypassable endokernel that is simple and guarantees isolation.

To achieve this the Endokernel enforces the following policies: secure loading and initialization so that all protection is configured appropriately; exports call gates for cross domain calls and ensures argument integrity and context-switching; inserts a monitor for all system calls so that they can be fully virtualized; monitors all address space and protection bit modifications to ensure isolation is not disabled; controls all signals so that they route through the endokernel before transitioning to any untrusted domain endoprocess; and handles concurrency to support multi-threaded execution.

**Interface** In the basic architecture, the endokernel transparently inserts itself and presents a minimal interface to the protected resources. All virtualized resources are accessed through calls into the endokernel. Access to address spaces and file systems are monitored through the system call interface. Other resources are memory based and since the untrusted domain has no access to the endokernel state, there is no need for an explicit interface. A endoprocess is created with an endokernel call that specifies a set of pages and entry points.

**Protection Sphere and Transitive Capabilities** Most operating systems provide a `fork` and `exec` functionality for creating new processes. If left unconsidered, as all prior work does, then the `fork`ed process will be able to access privileged state. Furthermore, on `exec` the runtime erases the process, which on first glance preserves the basic property, however, the new process could use the file system to access the parent’s process state. In this work we choose to ensure 1) all forks include the same endokernel that enforces the policy in the new address space, and 2) retain system interface restrictions in an `exec`ed process.

**Definition 3** (Protection Sphere). Each process defines an protection sphere, a context within which subspace isolation is guaranteed in and across all forked address spaces.

The implication is that a protection sphere provides a new realm of capability programming with the potential to tradeoff restricted runtimes for all forked processes (similar to work on Shill [48] and left for future exploration). To ensure subspace isolation, the following property must be maintained.

**Property 3** (Fork+Exec Transitivity). The endokernel transitive enforces the capabilities in each forked process.

## 4 Intravirt Design and Implementation

Intravirt is a userlevel only Endokernel system that fully virtualizes privilege and prevents bypass attacks. Beyond memory and CPU virtualization, it emphasizes full virtualization of system calls and signals, as well as exposes and addresses concurrency, multi-threaded, and multi-domain challenges. Intravirt injects the monitor into the application, as the trusted domain endokernel, and removes the ability of the untrusted domain to directly modify privileged state.

Privileged state includes: protection information (PKRU and memory mappings), code, endokernel code and data, direct system call invocation, raw signal handler data, CPU registers on transitions and control-flow, and system objects. The endokernel is inserted on startup by hooking all system call execution and initializing the protection state so that the trusted domain is isolated with no files opened or mapped.

### 4.1 Threat Model and Assumptions

We assume that an application is benign but potentially buggy: the same as prior work. We assume Intravirt is free of memory corruption bugs and trust the OS and hardware implementation of protection keys. An attacker can use bugs to launch a buffer overflow that both deviates control and injects a payload: a shell script or return, jump, or data oriented program. We assume an attacker can leverage any instructions within the memory space in an attempt to launch a shell, fork processes, and in general exercise any system interface. The attacker can attempt confused deputy attacks on Intravirt interfaces in order to launch TOCTOU and concurrency corruption. We do not consider architectural side-channels.

### 4.2 Privilege and Memory Virtualization

CPU and memory virtualization are the foundation of Endokernel isolation. Most of this work was generated by prior work in single-privilege level isolation (Nested Kernel, Erim, and Hodor). We provide a concise description here, and place a more complete overview is provided in Appendix A. Intravirt uses MPK protection keys to assign subspaces to endoprocesses. In our prototype, all pages of the trusted domain have key 0 and all pages of the untrusted domain have key 2. A virtual privilege switch occurs when the PKRU is modified so that access is granted to all domains, which is secured by using instruction capabilities to ensure the only WRPKRU instructions are inside secured call gates. To ensure these two are in place, we must ensure code, control, and memory integrity from system object, concurrency, and signal threats, which are detailed in subsequent sections. We add a new consideration for preventing processes from gaining control by switching into 32-bit compatibility mode, which changes how some instructions are decoded and executed. The security monitor code may not enforce the intended checks when executed in compatibility mode. Thus, we insert a short instruction sequence immediately after WRPKRU or XRSTOR instructions that will fault if the process is in compatibility mode. See Appendix D for more details.

### 4.3 System Call Monitor and Handling

Intravirt must ensure that access to system objects is virtualized. We could place this monitor in the kernel, however, that would separate the memory protection logic from the mechanism and create greater external dependencies. Furthermore, it would push the policy specification into the kernel, but the abstractions supported need to be extensible and thus will endanger the whole OS. Instead we observe that system resources are provided via `syscalls`, whose semantics are stable and allow for reasoning and enforcement.
of endoprocess isolation policies. Additionally, Intravirt will have greater portability if targeting POSIX. Even if we locate the monitor in the kernel, we will have to add extra context switches and layers of complexity in the kernel.

Intravirt virtualizes system objects by monitoring control transfers between the untrusted domain and the OS through a novel in-address space monitor, called the *nexpline*.

**Property 4 (Nexpoline).** All legitimate syscalls go through endokernel checks and virtualization.

The basic ways Intravirt does this include: 1) preventing all syscall operations from untrusted domain subspaces and 2) mediating and virtualizing all others. We could use a control-flow integrity monitor to provide both of these, like CPI [35], but that would add unnecessary overhead, require compiler support, and violate our minimal mechanism principle. Alternatively, we could extend the OS, but that would break our principle of no kernel dependencies and cost.

### 4.3.1 Passthrough

The first step is to determine what virtualization, if any, is necessary, because many syscalls do not allow endoprocess bypass. Additionally, a syscall will use the application’s virtual addresses to access the memory, which means that the MPK domain will be enforced even from supervisor mode—this is something we learned only through failing, so it is important to note that by default the kernel leaves the MPK domain untouched and thus the hardware continues to enforce the protection even from supervisor mode. The benefit of this is that any kernel access to endoprocess subspaces not permitted by the current PKRU value will trap into the endokernel—a powerful deny-by-default policy enforced even on ioctl s with unknown semantics. It does not mean that the kernel cannot remap pages and get around the domains, but a common path for access must be coded around, adding greater confidence to the protected access paths. With these passthrough syscalls, we use our protected nexpoline control path and right before executing the syscall, we transition the PKRU domain to the original caller so the kernel will respect the memory policies in place. After the syscall, Intravirt will switch to trusted domain to finalize the syscall and then transition back to the calling endoprocess.

### 4.3.2 No syscall from untrusted domain subspaces

To prevent direct invocation of syscalls, we remove all syscall instructions from untrusted domain and ensure integrity like we do for WRPKRU, however, the syscall opcode is short and might lead to high false positives. Instead, we use OS sandboxes that restrict syscall to a protected trusted domain subspace. There are two that can be used: seccomp and dispatch. When we started, seccomp was the only option, but has many drawbacks: 1) you cannot grow or modify a seccomp filter, making support for multi-threading and forks challenging; and 2) it adds significant overhead. The only way to address the second is to use a different mechanism. Thus, we explore a recently released kernel dispatch mechanism, a lightweight filter that restricts syscalls to the particular subspace. Both of these mechanisms work by specifying the virtual address region that is permitted to invoke system calls, which we use to restrict to endokernel subspaces.

### 4.3.3 Complete mediation for mapped syscall

Unfortunately, the only way to invoke a syscall is for the opcode to exist in the runtime, meaning it must be placed in the memory where the untrusted domain can jump to. Ideally, protection keys would distinguish executability and we could use a endoprocess switch, but they do not: Intel® relies on the NX mappings. Alternatively, subspaces with syscall opcodes could be marked NX, but the nexpoline would require another syscall to enable write access to the page.

Instead, the nexpoline protects each instance of syscall; return; instruction sequence, i.e., the *sysret-gadget*, so that if control neglects to enter through the call gate, the syscall is inaccessible. The basic control flow is to enter through the call gate and perform system virtualization, set up the nexpoline code subspace, jump to the syscall, then return to the handler for cleanup. We develop three isolation techniques: 1) randomizing the location of the *sysret-gadget* and restrict access to the endokernel, 2) making the instruction ephemeral by adding and removing it before and after each syscall, and 3) using Intel® CET hardware. These designs become complicated when considering concurrency, which we detail in §4.6.

**Randomized Location** To abuse the *sysret-gadget* the attacker must know its location, which is randomized in the first isolation approach. We create one pointer that points to the *sysret-gadget*, and make it readable by the endokernel endoprocess. This means that to get access to the pointer, the endoprocess must be switched-to first, and thus guarantee protected entry. The pointer is looked up immediately after switch, which means that all code between that instruction and the *sysret-gadget* will execute: endokernel executes all virtualization and once approved invokes the sysret-gadget. This ensures complete mediation because the only way to get the *sysret-gadget* location is to enter at the beginning. The *sysret-gadget* can then be re-randomized at various intervals to provide stronger or weaker security; we measure the cost of randomizing at differing numbers of syscalls.

**Ephemeral On-Demand** While randomization—especially if randomizing on each syscall—creates a high degree of separation, it is not guaranteed. To provide deterministic isolation, we present the ephemeral nexpoline, which achieves isolation by writing the *sysret-gadget* into an executable endokernel subspace on gate entry and rewriting to trap instructions (int 3) after completion. This requires Intravirt to create a single page for the nexpoline in the trusted domain with read and write permission restricted to the endokernel (via MPK) and execute permission for all domains. Intravirt ensures that when the untrusted domain executes, the entire page is filled with int 3 instructions, which would fault if the untrusted domain were to jump to this page. The endokernel interposes on all control transfers from the OS to untrusted domain, thus it ensures that prior to any control transfer back to the untrusted domain the *sysret-gadget* is removed. As a result, there is no
executable sysret-gadget while untrusted domain is in control.

**Control enforcement technology (CET)** [28] CET provides hardware to enforce control flow policies. While designed for enforcing Control-Flow Integrity [3], we show how to (ab)use CET to implement a virtual call gate, which ensures syscall; return; is not directly executable by the untrusted domain. Briefly, CET guarantees that all returns return to the caller and indirect jumps only target locations that are prefixed with the end-branch instruction. CET also supports legacy code, by exporting a bitmap to mark all pages that can bypass indirect jump enforcement, but the shadow stack must be used across the whole application.

Intravirt allocates a shadow stack for each endoprocess and ensures that a stack cannot be used by a different endoprocess by assigning each one to a protected subspace. Intravirt marks all endokernel entrypoints with ENDBR64: denying transitions into the endokernel from any indirect jumps. This creates a problem though, because indirect jumps within the endokernel also require end-branch instructions and could be used as alternative entrypoints to the endokernel. Thus, all jumps within the endokernel are direct jumps with a fixed offset from current IP and thus are not exploitable. This allows syscall; return; to be placed anywhere in the trusted domain, since the hardware automatically ensures all syscall will start from a legit entrypoint. While CET can provide greater security for the whole application, our evaluation shows significant overheads compared to the other approaches (see §6).

### 4.4 OS Object Virtualization

The primary goal of Intravirt is to preserve endoprocess isolation, which requires system object virtualization for eliminating cross endoprocess flows. Intravirt represents these in three three core system abstractions and policies to systematically reason about and specify policies: files (including sockets), address spaces, and processes.

**Sensitive but Unvirtualized Syscalls** A key class of system interfaces (ioctl, sendto, etc.) may index into regions of the address space that the kernel might accesses on behalf of a process, but as discussed, the kernel will use the userlevel virtual addresses which are protected by the hardware enforcing MPK domain isolation even from privilege accesses. These do not require full system level virtualization, but if the kernel did not implement that strategy, they could be fully virtualized by analyzing the arguments and denying any access that crosses endoprocess isolation.

**Files** The Linux kernel exposes (via the procs) several sensitive files that may leak endoprocess’s memory, because the kernel does not enforce page permissions, e.g., /proc/self/mems. [12]. To prevent any file-related system call from ever pointing to such a sensitive file, Intravirt tracks the inode of each opened file. Conveniently, inodes are the same even when using soft or hard links. This allows Intravirt to enforce that no open inode matches the inode of a sensitive file. The associated rules are transitively forwarded to child processes as they inherit the file descriptor table of the parent.

**Mappings** In addition, one may break the isolation property of Intravirt by aliasing the same file mapping multiple times with different access permissions. For instance one mapping may allow read/execute, while the other alias mapping to the same file permits read/write accesses. We prevent such attacks by emulating the mapping using the regular file interface and copying the file to a read/write page first which is later turned read/execute after all security checks passed. As a result an executable page is never backed by a mapped file.

**Processes [Remote Process Accessing]** The Kernel permits virtual memory accesses via process_vm_readv and process_vm_writev system calls. These calls access memory of remote processes or the current process itself. For these two system calls, we apply the same restrictions as for file-backed system calls preventing a domain from accessing another domain’s memory. In addition, we completely prevent access to another process’ memory via process_vm_readv/writev. [fork and vfork] Due to the insecure behavior of vfork, we emulate it by using fork instead. fork needs to be altered to enforce transitive policy enforcement across process boundaries. [exec] A process application can be modified using the exec system call. In this case, the kernel loads the new executable and starts executing it. This is problematic, because we need to initialize its protections before the application. Hence, any exec system call needs to be intercepted to ensure policy enforcement is enabled after exec.

**Forbidden system calls** Several syscall access protection state. Intravirt currently denies access to the following and leaves their virtualization to future work: clone with shared memory, pkey_* system calls, modify_ldt, rt_tgidgqueueinfo, seccomp, prctl accessing seccomp, shmdt, shmat, ptrace.

### 4.5 Signal virtualization

Signals modify process control flow by pushing a signal frame onto the stack and transferring control to the point indicated by signal handler. Beyond exposing the endoprocesses to control hijack, signals also expose the PKRU through struct sigframe, which would allow an attacker to modify policies. Intravirt virtualizes signals by adding a layer of indirection between untrusted domain code and the OS.

**Basic Operation** On signal registration, the endokernel tracks the handler in an internal table, i.e., the virtual signal, and then registers a real signal with the OS that points to a handler within the endokernel. Just like interrupt handling in a kernel, Intravirt splits signal handling into a top and bottom half. The bottom half receives a signal and sets up the top half to deliver it after performing a sigret to the kernel. Intravirt
only allows one signal to be queued for delivery at a time (by using signal masking), thus relying on the OS to manage nested signals without losing compatibility.

Concurrency and Reentry Signals can be delivered asynchronously—to the untrusted domain or the endokernel—or synchronously to the endokernel—while the monitor was virtualizing a syscall. In each case, Intravirt adds the signal to the pending queue and calls sigret, however, the handler sets up different return locations based on the endoprocess. If it is the untrusted domain, then Intravirt modifies the return location to be the exit gate so that all configurations are in place for untrusted domain execution. If in the endokernel, Intravirt returns to the interrupted state preceding the signal. This ensures that the monitor remains atomic and that all Intravirt virtualization is cleaned up prior to returning to the untrusted domain.

Default PKRU Domain Linux resets the PKRU to a default value on all signal delivery, which means it cannot deliver to the secure endokernel stack. We created a design and implementation to work with this limitation, but realized the signal handling interface should instead deliver the signal to the registering domain. Appendix F describes this design as it is far more complex to securely handle signals, but in our final prototype we decided to modify the Linux kernel to deliver signals to the registering PKRU domain and thus securely transition from the kernel to the endokernel.

4.6 Multi-threading and Concurrency

While concurrency is a well-known issue for multi-threaded monitors, prior systems have ignored its design and implementation. The issue is that the endoprocess abstraction allows concurrent threads and thus exposes memory to corruption that could modify execution through TOCTOU attacks.

Nexpoline Isolation with the Queen Thread An attacker could jump to the syscall-gadget when another thread is executing a syscall. The randomized nexpoline is simplest since we depend on probability and can enlarge the size of syscall-gadget region. CET is also simple by using per-thread shadow stacks to prevent unauthorized indirect jump. The most complicated design is Ephemeral nexpoline because it exposes the syscall-gadget with executable permissions. The solution is to create per thread syscall filters that restrict each thread to using a different page, so that even if another thread finds the location and calls it, the kernel will deny. Unfortunately, this design is complicated by the functionality of the filter. The dispatch mechanism is not a problem since it can reset the region on any new thread creation. However, seccomp inherits filters in all forks. Intravirt addresses this by introducing a special thread, called Queen, that has no seccomp filter and which spawns all new threads created by clone with a per thread nexpoline filter.

Monitor Atomicity TOCTOU attacks expose various state to corruption due to race conditions in endokernel processing. One solution is to use a monitor pattern that allows one thread to enter at a time, however, this would be prohibitively costly. Our solution is to use fine-grained locking based on the specific OS object being interacted with. First, we provide a lock for each file descriptor: each file descriptor can be concurrently accessed, but only one thread can access one file descriptor in one system call at a time. Second, for mapping based OS objects, we maintain a global lock that only one thread can call mapping based system calls. Third, we provide one global lock for the system calls managing signals.

4.7 Multi-Domain for Least Privilege

So far we discussed the abstraction of a single untrusted endoprocess and a privileged endokernel. In this section we extend the abstraction to allow multiple untrusted endoprocesses. This allows developers to decompose applications into least-privilege functional units. While Intel® MPK is limited to 16 domains in total (and 14 usable for Intravirt), the design is independent of the number of domains and supports virtualizing of domains as suggested by libMPK [54]. Intravirt builds on MPK to provide an interface to isolate code and data, create a endoprocess, switch endoprocesses via xcall, and a utility library to assist the needs of a endoprocess.

Secure Dynamic Creation of endoprocesses We overload the existing endokernel system call monitor to handle a virtual system call iv_create_domain(code, data, entrypoints, xcall_stub). Intravirt assigns an unused MPK domain to the endoprocess, and maps the code and data to this domain to prevent other domains from accessing them. It additionally virtualizes OS objects and signals as described in § 4.4 and § 4.5 and monitors the system calls of all endoprocesses to prevent privilege escalation. Switching into the newly created endoprocess is restricted to the provided entrypoints. To improve the switching performance, Intravirt installs a secure domain switch at the xcall_stub which is used to call other endoprocesses.

Securely Switching between endoprocesses via xcall Switching between endoprocesses is supported via xcall (endoprocess, entrypoint_id, ...). Each switch is mediated via the endokernel which holds the information for the entrypoints of each endoprocess. Therefore, xcall switches to endokernel and looks up the entrypoint of the target endoprocess without destroying the callee arguments. It then updates the currently running endoprocess and switches to that endoprocess (including a stack switch) to proceed to the function call. During that process signal delivery to the current endoprocess is disabled to prevent leakage via signals.

4.8 Implementation Details

Intravirt is built out of five primary components—secure loading, privilege and memory virtualization, syscall virtualization, signal virtualization, and xcall gates. We use the Graphene passthrough LibOS [71–73] to securely load, insert syscall hooking into glibc, and separate the trusted domain from untrusted domain memory regions. We use ERIM [75] to isolate memory and protect WRPKRU, and 200 LoCs for tracking page attributes. We implement all syscall and signal virtualization code. In total our system comprises ≈15k lines of code, with ≈6,400 new Intravirt code.
5 Programmable Least-Authority

We demonstrate the extensibility of the Endokernel Architecture by developing virtual privilege rings and using them to 1) eliminate the recent sudo vulnerability [15] and 2) sandbox buggy parsers and isolate sensitive OpenSSL data in NGINX.

5.1 Separation Facilities: Nested Boxing

Least-authority is hard to apply in practice because security policies are highly dependent on the objects being protected. As indicated, many abstractions are rigid and do not allow for specialization from the application developers. The Endokernel Architecture allows the ability to use the endokernel to explore diverse endoprocess and sharing models. Thus, we present the nested boxing abstraction, which effectively creates virtual privilege rings in the process. The nested boxing model allows each level access to all resources of the less privileged layers, while removing the ability from those domains to access more privileged domains. In this paper we fix the number of domains to four from most to least privileged: endokernel, safebox, unbox, and sandbox. Each domain is given an initialized endoprocess that provides capabilities for accessing domain resources. To make programming easier, we also use a libos that aids in allocation and separation policy management.

Dynamic Memory Management One of the core challenges with privilege separation is modifying the code so that data is statically and dynamically separated. Static separation is easily done using loader modification, but dynamic memory management is harder, in particular when we have to ensure subspace isolation. In our system we provide a nested endokernel allocator that transparently replaces whatever allocator the code originally used and automatically manages the heap and associated privilege policies.

Memory Sharing An endoprocess shares data through a simple manual page level grant/revoke model. An endoprocess grants access to any of its pages to a lower privilege domain and removes access through the revoke operation.

Protected Entry and Return Cross domain calls, or xcalls, are invoked by the calling domain and can only enter the called domain at predefined entry points as specified by the endoprocess definition. This interface will reject all attempts of accessing the safebox if it is not to a preloaded entry point. It will then do the domain-switch: switch the stack, current domain ID, store the return address in a protected memory subspace, and transfer control to the safebox. When the called function finishes, it returns to the interface function, which domain-switches back to the untrusted domain. Entry points can either be defined manually or as we show for full library separation, by using the library export list. This model of control flow allows the called domain to subsequently call less privileged code, i.e., if it does this the called code operates within the endoprocess context and is thus in the TCB. We allow users to determine when and how to use these features, granting greater flexibility at the cost of more complexity in reasoning about security if a callback is issued. This can implement the Shreds abstraction, if used in code with no callbacks.

5.2 LibOS and Automated Transformation

Programming the low-level interface affords power at the cost of effort required to use it. To simplify use we use a LibOS, called libsep, to aid developers in using the sandbox and safebox isolation, and we also provide an automation tool that separates at library boundaries.

libsep provides a simple interface for developers to statically annotate their code so that all data, code, and entry points are automatically separated during loading. Our approach uses section labels to do this. Additionally, the libsep tools automatically align subspaces for encapsulation and generates the stub functions. Instead of modifying the code with xcalls, entrypoints are automatically wrapped to invoke the xcall on use. Last, libsep provides a simple thread safe slab allocator.

Whole library isolation Alternatively Intravirt can separate entire libraries. The developer adds a function that specifies the library virtual address and then the libos labels all exported symbols as entry points and generates stub code. A key element of this design is that domain transitions only occur in one direction because we use it as a safebox. So if a library has a callback into a sandbox or unbox, that code will operate in the safebox privilege level.

5.3 Use Cases

In this section we detail two use cases of nested boxing.

5.3.1 Eliminating sudo Privilege Escalation

A recent bug was found in the sudo argument parser that allows an attacker to corrupt a function pointer and gain control with root access [15]. We compartmentalize sudo so that the parser code, in file parse_args.c, is sandboxed, and restricted to only the command line arguments and an output buffer. The worst attack that can happen now is overflowing its internal buffer and eventually segfault or done nothing harmful. In summary, by changing approximately 200 lines of codes, importing our libsep in sudo and using Intravirt, we confine the argument parser and successfully prevent the root exploit. More generally, almost all parsers have a similar type of behavior and could benefit from similar changes, and possibly automatically.

5.3.2 Towards a Least-Privilege NGINX

We present a novel compartmentalization of NGINX, which allows us to measure the effort required and performance costs of improving the security of a complex system while using the nested boxing facilities. Our aim is to sandbox the NGINX parser and safebox the OpenSSL library. The value of the sandbox is that it defends against common buggy entry points (e.g., CVE-2009-2629, CVE-2013-2028, and CVE-2013-2070) used by attackers to launch control-flow hijacks (by injecting HTTP requests) and leaking or modifying sensitive state (e.g., cookies, another client’s responses, or keys). The value of the safebox is to eliminate leakage of keys to the majority of the application, a TCB reduction argument, including session and private keys.

Safeboxing OpenSSL To prevent key leaks, we safebox the entire OpenSSL libraries using libsep, which links a special loader function that identifies all OpenSSL code
and data sections during the start up time, and links libseep allocators, and initializes the endprocess context. The tool also identifies entry points from the exports and instruments call gates for domain transitions. The developer minimally needs to link libseep into the application, identify the addresses of the sections, and call a new custom system call defined in Intravirt, then Intravirt does the rest—it can easily be applied to isolate other libraries. While this approach protects against vulnerabilities outside of OpenSSL, it does not prevent attacks from a compromised OpenSSL like Heartbleed [18]. To prevent such attacks, we could split OpenSSL similar to ERIM [75]. Another key element is that our abstraction only instruments calls into the library, so if the library calls back into the unboxed domain, then it will be in the TCB.

**Sandboxing HTTP Parser** The Nginx parser is because its functionality is limited while it has been historically buggy and grossly overprivileged. While the parser only interacts with a small subset of data, that data is referenced from a large structure, making sharing challenging. We address this by allocating each instance of the structure into a special page that the sandbox is granted permission to access upon invoking the parser and permission is revoked once finished. This allows us to minimize overprivilege to only the per request data structure and only for a short period of time, representing least-privilege.

We manually modified NGINX HTTP request handler to identify the address of the parser functions, and install the call gates by calling custom system call in Intravirt. As well, we aligned the data structures with the page sizes to properly work with MPK and we granted the data structures to the sandbox.

## 6 Evaluation

In this section we evaluate the security and performance of Intravirt. We highlight Intravirt’s security properties protecting against known [12] and additional attacks found by us. Subsequently, we investigate the performance characteristics of Intravirt in several microbenchmarks, on common applications, and least-privilege Nginx use case.

### 6.1 Security Evaluation

Table 1 summarizes the quantitative security analysis based on known attacks described by Conner et al. [12] and additional attacks found by us. Intravirt defends against the attacks raised in [12] and new attacks we found by virtualizing and monitoring OS objects and preventing privilege escalation via signals, multi-threading, multi-domain. We briefly discuss new attack vectors and refer the reader for details to the Appendix B. The attacks try to bypass Intravirt by performing system calls modifying the protection policies or try to elevate privileges by overriding the PKRU register.

| Attack                                      | secc_rand | secc_eph | CET |
|---------------------------------------------|-----------|----------|-----|
| Inconsistency of PKU Permission [12]        |           |          | ★  |
| Inconsistency of PT Permissions [12]        |           |          | ★  |
| Mappings with Mutable Backings [12]         |          ★ |          | ★  |
| Changing Code by Relocation [12]            |          ★ |          | ★  |
| Modifying PKRU via sigreturn [12]           | ★         |          | ★  |
| Race condition in Signal Delivery [12]      |          ★ |          | ★  |
| Race condition in Scanning [12]             |          ★ |          | ★  |
| Determination of Trusted Mappings [12]      |          ★ |          | ★  |
| Influencing Behavior with seccomp [12]      |          ★ |          | ★  |
| Modifying Trusted Mappings [12]             |          ★ |          | ★  |
| Forged Signal                               |          ★ |          | ★  |
| Fork Bomb                                   |          ★ |          | ★  |
| Syscall Arguments Abuse                     |          ★ |          | ★  |
| TSX attack                                  |          ★ |          | ★  |
| Race condition                             |          ★ |          | ★  |

Table 1: Quantitative security analysis based on attacks demonstrated in [12] and attacks found by us. ★ indicates the variant of Intravirt in this column is vulnerable, ◦ if it prevents this attack.

Infinite amount of processes trying to make a system call by guessing the nexpoline’s location. This bypasses the regular entrypoint of the nexpoline and any security checks. The ephemeral and CET nexpoline are not vulnerable to this attack.

**Syscall Argument Abuse** An attacker may use pointers in unvirtualized system call that point to memory of a privileged domain to test for bits in the contents.

**TSX Attack** Intel® TSX [27] supports transactional memory which performs rollback of memory and execution. This feature can be used to efficiently probe memory contents. Only the random nexpoline is vulnerable to this attack, since the secret random location can be probed via TSX.

**Race Conditions** Race conditions in Intravirt could be exploited in multi-threaded and -processed applications. Shared memory across multi-processes provides a potential avenue to modify system call arguments by a second process after Intravirt performed the necessary security checks. To prevent these multi-process attacks, Intravirt relies on its own copy of the system call arguments which cannot be access by another process. Similar attacks may also be performed by multiple threads. In particular, the interleaving of system calls can be exploited to gain access to secrets.

## 6.2 Performance

In this section we characterize the performance overhead of Intravirt. First, we explore microbenchmarks focussing on the cost to intercept system calls and signals. Second, we demonstrate the performance of Intravirt for common applications. Third, we evaluate the cost of the least-privilege Nginx use case.

We perform all experiments on an Intel® 11th generation CPU (i7-1165G7) with 4 cores at 2.8GHz (Turbo Boost and hyper-threading disabled), 16GB memory running Ubuntu 20.04, and the kernel version 5.9.8 with CET and syscall dispatch support. For all experiments we average over 100 repetitions and analyze different Intravirt configurations. Intravirt
relies on a Seccomp filter or a syscall user dispatch (denoted by Sec or Dis) for system call interception, and random, ephemeral, or CET trampoline (denoted as rnd, emp, cet). In this configuration space we evaluate 5 different configurations ((secc/disp)|rand|eph|cet) excluding disp — rand configuration because it’s major overhead is from the random number generation. Throughout this section, we compare against MBOX [31] and strace, ptrace-based system call monitors. MBOX fails for experiments using common applications. In these cases we approximate the performance of MBOX using strace. In our microbenchmarks strace outperforms MBOX by 2.7% providing a conservative lower bound for MBOX.

6.2.1 Microbenchmarks

System call overhead We evaluate Intravirt’s overhead on system calls and signal delivery in comparison to native and the ptrace-based techniques. Figure 3 depicts the latency of LMbench v2.5 [45] for common system calls. Each Intravirt configuration and the ptrace-based techniques intercept system calls and provide a virtualized environment to LMbench while protecting its privileged state.

secc_eph and secc_rand _1 modify the trampoline on every system call, but secc_eph saves the cost of randomizing the trampoline location and hence, incurs less overhead. secc_eph/disp_eph, and secc_cet/disp_cet demonstrate the performance difference between using a Seccomp filter or syscall user dispatch to intercept system call invocations. Overall, disp_eph outperforms all other configurations, while secc_rand _1 is the slowest. Even though CET relies on hardware support, it does not outperform other configurations. Intravirt adds 0.5 - 2 usec per system call for disp_eph for policy enforcement and domain switches. In comparison the ptrace-based technique incurs about 20 usec per invocation which is 4.7-26.8 times slower than disp_eph. We observe high overheads for Intravirt protecting fast system calls like read or write 1 byte (126%-900%), whereas long lasting system calls like open or mmap only observe 29%-150% overhead. We demonstrate the difference by performing a throughput file IO experiment. Figure 4 shows high overheads for reading small buffer sizes which amortize with larger buffer sizes. Since overhead induced by Intravirt is per syscall basis, to read a file with bigger buffer size has much less overhead than with the smaller buffer size. Even though we observe high overheads for some system calls, applications infrequently use them and observe far less overhead as shown for common applications in § 6.2.2.

Randomization and performance tradeoff The secc_rand configuration rerandomizes the trampoline for each system call generating an random number using RDRAND (approx. 460 cycles). We explore alternative randomization frequencies to amortize the cost of randomizing over several system calls. We tradeoff performance with security, since the system call address is simpler to guess if randomization happens less frequently. The goal is to find a reasonably secure, but fast randomization frequency. Figure 5 evaluates getppid system call for different randomization frequencies. getppid is the fastest system call and hence, results in the highest overhead of Intravirt. The overhead of secc_rand amortizes with less frequent randomization and does not improve much beyond 16 system calls per randomization. secc_rand at 4 system calls per randomization shows similar performance with secc_eph’s performance which we also observed for other LMbench microbenchmarks.

Thread scalability To prevent race conditions and T0C-T0U attacks in Intravirt, locks protect Intravirt’s policy enforcement as addressed in the § 4.6. We demonstrate the scalability of Intravirt in figure 6 using the sysbench [33] tool which concurrently reads a 1 GB file from varying number of threads. Due to the additional locks in Intravirt, the number of futures system calls increases with the number of threads.

At 4 threads all CPU cores are busy and we observe the best performance. The overhead of each configuration is similar to the microbenchmarks. secc_cet and disp_cet suffer a performance decrease of up to 60%, because the syscall performance of CET-based configurations is the lowest. Compared to strace, Intravirt outperforms by 4.3-8.2 times.

6.2.2 Overhead on Applications

Along with the microbenchmarks, we analyze the performance of common applications such as lighttpd [38], Nginx [52], curl [14], SQLite database [66], and zip [26] protected by Intravirt. Figure 7 shows the overall overhead of each application compared to the native execution.

curl [14] downloads a 1 GB file from a local web server. It is particularly challenging workload for Intravirt, since curl makes a system call for every 8 KB and frequently installs signal handlers. In total it calls more than 130,000 write system calls and more than 30,000 rt_sigaction() system calls to download a 1 GB file. However, libcurl supports an option not to use signal, which reduces the overhead about 10% in average for Intravirt but strace gets worse about 140%.

Lighttpd [38] and Nginx [52] serve a 64 KB file requested 1,000 times by an apachebench tool [1] client on the same machine. All configurations perform within 94% of native. disp_eph outperforms all other configurations and highlights Intravirt’s ability to protect applications at near-zero cost with a throughput degradation of 1%. In contrast, strace has about 30% overhead.

SQLite [66] runs its speedtest benchmark [66] and performs read() and write() system calls with very small buffer size to serve individual SQL requests. Contrary to the microbenchmarks, difference between configurations is larger. Configurations using syscall user dispatch (disp_eph and disp_cet) observe about 30% less overhead when compared to their Seccomp alternatives (secc_eph and secc_cet). Strace performs poorly at more than 500% overhead.

zip [26] compresses the full Linux kernel 5.9.8 source tree, a massive task which opens all files in the source tree, reads their contents, compresses them, and archives them into a zip file. The observed performance degradation is in-line with the microbenchmarks for openat(), read(), and write() system calls.

Summary: Network-based applications like lighttpd and Nginx perform close to native results whereas file-based
applications observe overheads between 4 and 55% depending on the test scenario. Most impacted are applications which access small files like SQLite. In comparison to ptrace-based techniques, Intravirt outperforms by 38-529%.

6.2.3 Least-Privilege Nginx Performance

We evaluate the least-privilege Nginx shielding OpenSSL and sandboxing the HTTP parser (see § 5.3.2). We measure the throughput downloading files with varying sizes and normalize to the native performance (see Figure 8). Strace suffers from its interception costs and falls below 50% for large files. The results indicate less than 10% overhead for the different nexpoline techniques and the lowest for disp_eph at 3-5%. The number of system calls scales with file size causing decreasing the performance of Intravirt. Each xcall consumes 269 cycles for CET-based nexpoline (see Appendix C) and 116 CPU cycles for random and ephemeral.

### 7 Related Work

Traditional address space isolation provided by the operating systems is hardly used for least privilege use cases. Its
we chose Intel SFI [76] started a research field to inline security checks. To the contrary, Intravirt takes advantage of intra-process isolation which avoid these switches. Their approaches can be classified into language-based, OS-based and hardware-based techniques.

Language-based techniques Software Fault Isolation (SFI) [76] started a research field to inline security checks within application code. The goal is to translate an application via a compilation pass to enforce security properties such as Control-Flow Integrity (CFI) [2], Code-Pointer Integrity [34], or intra-process isolation [32]. Early techniques such as CCured [50] and Cyclone [30] insert bounds checks at compile time to trigger exceptions during the execution. These techniques have, e.g., been used to protect Java applications from Java Native Interface [11, 64, 67, 69]. Similarly, web applications need protections from the native code [81] which has lead to the development of Webassembly. Webassembly has become the defacto standard for language-based intra-process isolation. These techniques suffer from non-negligible performance overheads and similarly to Intravirt have to also protect against attack vectors via the OS. Intravirt provides the Endokernel Architecture to generalize this abstraction and could also be combined with language-based techniques, but we chose Intel® MPK for its superior performance.

OS-based techniques To improve upon the limitations of existing process-based isolation, several kernel abstractions have been suggested. light-weight contexts(lwC) [39], secure memory views(SMV) [25] and nested kernel [16] introduce a light-weight form of processes allowing to associate multiple virtual address spaces with the same process allowing for faster switches. Shreds [10] combines OS-based and language-based techniques to further improve the performance, and Wedge [7] combines with runtime checks. While these techniques improve the state of the art, their abilities are limited by the provided hardware performance for privilege level switching. To the contrary, Intravirt takes advantage of intra-process isolation which avoid these switches.

Hardware-based In recent years CPU vendors have suggested techniques to bypass the costly privilege level switch by, e.g., providing a VMFUNC instruction [29] to switch address spaces in the CPU from userspace or alter memory access permissions via userspace instructions like WRPKRU. Dune [6] uses Intel® VT-x to elevate a process to run both in privileged and unprivileged CPU mode. This allows to use the system call boundary to efficiently decompose applications and isolate components. Secage [40] achieves a similar separation using Intel®’s VMFUNC instruction which allows userspace applications to switch into a different address space. Koning et al. [32] build a tool to compile applications with varying isolation techniques and demonstrate the performance differences. Generally, the overhead is dominated by the efficiency of the memory permission switch technology or the language-based runtime security checks. ERIM [75], HODOR [24] and Donky [62] enable the secure use of Intel® MPK for memory isolation and demonstrate an efficient isolation technique. Unfortunately, their techniques have several short comings as demonstrated by Conner et al. [12] which Intravirt addresses without revert- ing to inefficient isolation techniques keeping the premise of MPK-based isolation techniques. The general applicability of MPK-based techniques is demonstrated by FlexOS [37] and Sung et al. [68] which use MPK in a unikernel environment.

System Call and Signal Virtualization Pitfall [12] demonstrated how the operating system can be used to bypass intra-process isolation as discussed in the previous section. To prevent such bypass, all operating system interactions need to be mediated, strict security policies enforced, and then virtualized. This style of attack elevated the importance of system call monitoring and virtualization while also demonstrating that existing kernel-based techniques are insufficient. Linux security module(LSM) [79] intercept system calls in the kernel and can be used to implement a system call filter as suggested by SELinux [41], AppArmor [5], Tomoyo [23] and Smack [61]. However, LSM would have to be extended to recognize different domains in userspace and its interface does not support modifying system call arguments which makes virtualization infeasible. Alternatively, Seccomp [13] offers the userspace a programmable system call filter using eBPF programs. Seccomp is widely used by numerous applications such as Native Client [81] to limit the OS interface accessible to the untrusted component to a bare minimum. Seccomp cannot easily be extended to filter different domains. Intravirt in some configurations relies on Seccomp and works around its limitations.

To virtualize the system calls, there are a number of efforts utilizing ptrace() system call. ptrace() allows a process to attach another process and monitor the memory and system calls and it is widely used for the debugging and the profiling purpose like strace utility. ERIM [75] use ptrace to intercept memory-related system calls such as mmap() and virtualize the system calls to fulfill with the memory policy. However, ptrace() approach is known to suffer from the high performance overhead because the system call interposition requires multiple context switches.

8 Conclusion

This paper introduces Endokernel Architecture, which builds a new virtual machine abstraction for representing subprocess...
authority. A self-isolating monitor efficiently enforces the abstraction by mapping the subprocess authority to system level objects. We demonstrate the usefulness of Endokernel Architecture by decomposing Nginx to enforce least-privilege for its HTTP parser and OpenSSL library. Our prototype, Intravirt, explores implementation tradeoffs and shows that sub-process authority can be enforced efficiently (within 5% of native performance).

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A Intra-Process Memory Isolation

This appendix describes the background of intra-process memory isolation, including the abstractions and pitfalls of memory protection keys, and the systematic abstractions needed for a comprehensive memory isolation mechanism.

A.1 Protection Background

Hsu et al. [25] describe three generations of privilege separation, each increasing from manual, address-space isolation to the third generation that efficiently enables concurrent per-thread memory views. The key is new hardware that extends paging with userlevel tags for fast but insecure isolation.

Intel® Memory Protection Keys (MPK) MPK [29] extends page tables with a 4-bit tag for labeling each mapping. A new 32-bit CPU register, called PKRU, specifies access control policies for each tag, 2-bits per for controlling read or write access to one of the 16 tag values. The policy is updated via a new ring-3 instruction called WRPKRU. On each access, the CPU checks the access control policy as specified by the mapping’s tag and associated policy from the PKRU. If not permitted the CPU faults and delivers an exception.

MPK Security vs Performance Unfortunately, the PKRU can be modified by any userlevel WRPKRU instruction: MPK is bypassable using gadget based attacks. As such MPK balances security and performance by allowing protection changes without switching into the kernel.

Preventing MPK Policy Corruption Nested privilege separation reconciles the exposure of protection state by ensuring WRPKRU instructions are only used safely by the endokernel. They achieve this by removing all WRPKRU instructions from the untrusted binary and crafting nested call gates that prevent abuse [10, 22, 24, 37, 68, 75].

A.2 Privilege and Memory Virtualization

In this section, we provide an overview of ERIM [75], which we extend to build Intravirt. We encourage the reader to review detailed methodology from the original work.

Thread Model An initial configuration partitions the application into the trusted domain and untrusted domain, where the trusted domain contains the trusted monitor and the untrusted domain contains the rest. Once the application is separated so that the parts are differentiated, the system is configured so that all pages of the trusted domain have key 0 and 1 based on the confidential requirement, and all pages of the untrusted domain have key 2. Some pages will have other keys if they belong to other subdomains in untrusted domain.

Virtual Privilege Switch One of the most important elements when nesting the endokernel into the same address space is the need for secure context switching, which is complex to get correct because an attacker has access to
whatever is mapped into the address space. While the trusted domain is executing the PKRU is configured to allow_all (read/write to all domains), and operating in the trusted domain virtual privileged mode. While the untrusted domain is executing the PKRU value for the key 0 will be deny_write and deny_all for 1. The virtual domain switch is implemented as a change in the protection policies in the PKRU—when entering the monitor set the policy to allow_all, when exiting restore the original key based on the previous state. This means that whenever the value of PKRU changes so too does the currently executing domain. Each entry point into Intravirt is setup with a call gate with a WRPKRU that transitions the domain. The basic idea is to nest monitor code directly into the address space of the application and wrap each entry and exit point with a WRPKRU operation. By doing this the system can transition between contexts and only allow monitor code to access protected state—a virtual privilege switch. This similar technique is also used to switch between different subdomains to enable the usage of other keys in untrusted domain.

Securing the Domain Switch Unlike systems with real hardware gates, this SW/HW virtual privilege switch has challenges because the instruction must be mapped as executable to allow fast privilege switching. The first thing an attacker could do is use a direct jump to any code in the monitor and thus bypass the entry gate. This would in fact allow the attacker to execute monitor code. One way to thwart could be to modify the executable policy on the monitor pages, but that would require a call into the OS which defeats the purpose of fast domain switching of MPK in the first place. Instead, we observe that even if an attacker is able to jump into the middle of the monitor the domain would have never switched, therefore, none of the protected state is available for access and therefore the basic memory protection property holds. The only way to change the domain is to enter through the entry gate.

Since the switch is a single instruction, we can easily verify the result of such switching immediately after the WRPKRU instruction and loop back if it is not switch to the intended PKRU state. This ensure that the PKRU state at all exits of the gate sequence will be the intended PKRU state.

Effectively, the attacker now faces the dilemma that jumping into the middle of the code will ended nothing since it is the equivalent of running the same code in any other locations, or it can try to jump to the entry gate, but any landing places of the gate will only switch to the correct PKRU value and continue the execution with deterministic control flow. No code can be abused.

Instruction Capabilities Alternatively, an attacker could generate their own unprotected variant of WRPKRU—if an attacker can inject or abuse the WRPKRU instruction, they could switch domains and gain access to the monitors protected state. To deal with this ERIM and others like it use a technique called instruction capabilities: that is by using a combination of static transformations and code validation and dynamic protections an instruction becomes much like a capability. The static analysis removes all instances of the WRPKRU opcode so that the attacker has no aligned or unaligned instructions that could write the value without monitoring, and dynamic runtime is configured so that all code is writeable or executable but not both.

Controlling mode switches Processes may switch into 32-bit compatibility mode, which changes how some instructions are decoded and executed. The security monitor code may not enforce the intended checks when executed in compatibility mode. Thus, we insert a short instruction sequence immediately after WRPKRU or XRSTOR instructions that will fault if the process is in compatibility mode. See Appendix D for more details.

B New Attack Vectors and Security Evaluation

In addition to the attacks described by Conner et al. [12], we found several attacks against intra-process system call and signal virtualization. For the evaluation, we created a fixed address secret inside trusted domain. All test cases try to steal this secret and hence, would break Intravirt’s isolation guarantees. The attacks try to bypass Intravirt by performing system calls modifying the protection policy of the secret or trying to elevate itself to be trusted by overriding the PKRU register. They specifically target the implementation of Intravirt and highlight the degree to which Intravirt has followed through with its security guarantees. Table 1 summarizes the results.

Forged Signal Intravirt effectively prevents the basic sigreturn attack from [12]. However, the kernel places signals on the untrusted stack and delivers the signal to our monitor signal entrypoint. The untrusted application may forge a signal frame and directly call the monitor’s signal entrypoint. As a result, it can, e.g., choose the PKRU value and the return address. Therefore, the entrypoint has to distinguish between a fake signal from the untrusted application or a real signal from the kernel. The entrypoint is carefully constructed such that a signal returning from kernel returns with privileges from the masked trusted domain and hence, is capable of writing trusted memory. We rely on this observation and place an instruction at the beginning of the monitor which raises a flag in the trusted monitor. Any fake signal created by the untrusted application cannot raise the signal flag in trusted memory which violates a check that cannot be bypassed in the monitor’s signal entrypoint.

Fork Bomb This attack targets the random location of the system call instruction in Intravirt. To perform a system call the untrusted application may guess the random location of the system call instruction. Assuming the trampoline size is 16 pages, there are 65534 possible locations of the system call instructions. When the untrusted application is capable to fork children, the untrusted application may try different locations within each child. In case the child crashes, the system call was unsuccessful and the untrusted application has to retry. Using this brute force algorithm the untrusted application tries until a child does not crash. At this point the untrusted application has access to a child process that bypassed Intravirt’s security guarantees and may perform arbitrary system calls. It should be noted that only secc_rand is successively to this attack, since secc_eph removes the system call instruction completely when returning control to the untrusted application.
Syscall Arguments Abuse  Intravirt virtualizes a subset of all system calls. System calls which are not virtualized could be exploited to read secret memory, unless Intravirt verifies that all pointers provided to a system call lie within untrusted memory. We perform an attack based on the rename system call and pass it a memory pointer from the trusted domain as an argument. Intravirt successfully prevents this attack by checking the pointer locations.

Race Condition in Shared Memory Access  Shared memory may be used across multiple processes to bypass Intravirt checks on arguments to system calls. In particular, we consider a `pwritev`-based attack in which a child process performs a `pwritev` system call using an IO vector in shared memory. If the parent was permitted access to the same shared memory, it could alter the IO vector’s values to point to trusted memory. This attack is efficient such that the child’s monitor has already performed the security checks, but the system call has not yet read the affected IO memory vector. Intravirt prevents such attacks by copying pointers in system call arguments to the trusted memory region and only then performing the system call using the copied arguments.

Race Condition in Multi-Threaded Applications  Supporting multi-threading is essential in modern computing environment that Intravirt also supports it. But, there are a few attack surfaces which use race conditions in multi threading environment. First, indirect jump to `syscall; return;` is possible in ephemeral Intravirt. For example, one thread calls a `syscall` which take very long time, and the attacker thread jumps to the active `syscall; return;`. To prevent such attacks, we use either Syscall User Dispatch, or per-thread Sexecmp filter. Second, the attackers could perform TOCTOU attacks in the `syscall` virtualization. For example, one thread open a normal file and call a file-backed `syscall`, while another thread close the file descriptor and open a sensitive file which is not allowed for the untrusted code. In Intravirt, we provide locks per file descriptor that `close` system call could be locked when another thread is using that file descriptor. As well, Intravirt provides a lock for memory management `syscalls` and signal related `syscalls`.

TSX Attack  TSX is an extension to support transactional memory in x86. It has a similar principle as exception handling but at the hardware level. When any considered as a violation of transactional happens, the hardware rollback and modification jump to a preset restore code. Unfortunately, because of the rollback feature provides a harmless way of content probing since the first introduction, it has been used as a source of memory leakage. It has been obsoleted in the latest Intel® CPU but still exists in many products with MPK. Our attack utilizes TSX as a probe to the randomized trampoline. First, a `xbegin` is used to enable the TSX environment. Then, we call an address within the trampoline region. Now, there are three cases about the content on target address, `int3`, `syscall` and `ret`. For the first two cases, TSX will be aborted but in the second case, `ret` instruction can be executed successfully. Such difference is sensible from the view of the attacker and the address contains `ret` is exposed. Because our syscall gadget is `syscall; ret`. This exposed the secret address of `syscall`. Fortunately, TSX can be disabled through kernel or BIOS, and among all Intravirt configurations, only `secc_rand` is secret-based and susceptible.

C  Analyzing CET

The experimental results for CET indicated much higher overheads than we anticipated, and so we explored CET under different scenarios to understand if the overheads were truly CET or if there might be something in the Linux implementation or Intravirt. To understand the details we perform a minimal experiment on LMbench, Lighttpd, and Nginx.

We report on three configurations: 1) no CET at all as baseline, 2) CET is enabled in the benchmark application binary and the dependent shared libraries, and 3) CET are enabled all the dependent binaries including glibc.

The result of the tests says that both CET enabled configurations have 2-8% of overhead compare to the baseline configuration, that is, only enabling CET induces at most 8% overhead. In addition, IntraVirt performs multiple domain management, which contains multiple shadow stack management and shadow stack switch on domain switch, therefore CET enabled version of IntraVirt has more overhead than others.

Examining CET performance is a key element of related work but was out of scope for the analysis at this time.

D  Controlling mode switches

This appendix explains why it is necessary to check that the application is executing in 64-bit mode when it enters the trusted code, and it also describes a mechanism for performing that check.

64-bit processes on Linux are able to switch to compatibility mode, e.g. by performing a far jump to a 32-bit code segment that is included in the Global Descriptor Table (GDT). Executing code in compatibility mode can change the semantics of that code compared to running it in 64-bit mode. For example, the REX prefixes that are used to select a 64-bit register operand size and to index the expanded register file in 64-bit mode are interpreted as `INC` and `DEC` instructions in compatibility mode. Another example is that the RIP-relative addressing mode in 64-bit mode is interpreted as specifying an absolute displacement in compatibility mode.

Executing the trusted code in compatibility mode may undermine its intended operation in a way that leads to security vulnerabilities. For example, if the trusted code attempts to load internal state using a RIP-relative data access, that will be executed in compatibility mode as an access to an absolute displacement. The untrusted code may have control over the contents of memory at that displacement, depending on the memory layout of the program. This may lead to the trusted code making access control decisions based on forged data. Conversely, if the trusted code stores sensitive data using a RIP-relative data access, executing the store in compatibility mode may cause the data to be stored to a memory region that can be accessed by the untrusted code.

To check that the program is executing in 64-bit mode when it enters the trusted code, a sequence of instructions such as the following may be used:
1. Shift RAX left by 1 bit. In compatibility mode, this is executed as a decrement of EAX followed by a 1-bit left shift of EAX.
2. Increment RAX, which sets the least-significant bit of RAX. In compatibility mode, this first decrements EAX and then increments EAX, resulting in no net change to the value of EAX.
3. Execute a BT (bit test) instruction referencing the least-significant bit of EAX, which is valid in both 64-bit mode as well as compatibility mode. The BT instruction updates CF, the carry flag, to match the value of the specified bit. It does not affect the value of EAX.
4. Execute a JC instruction that will jump past the next instruction iff CF is set.
5. Include a UD2 instruction that will unconditionally generate an invalid opcode exception, which will provide an opportunity for the OS to terminate the application. The security monitor should prevent the untrusted code from intercepting any signal generated due to invalid opcode exception from this code sequence.
6. Shift RAX right by 1 bit to restore its original value. This instruction is unreachable in compatibility mode.

The preceding description of the operation of the instructions in compatibility mode assumes that the default operand size is set to 32 bits. However, a program may use the modify_1dt system call to install a code segment with a default operand size of 16 bits. That would cause the instructions that are described above as accessing EAX to instead access AX. That still results in the instruction sequence detecting that the program is not executing in 64-bit mode and generating an invalid opcode exception. Furthermore, Intravirt can block the use of modify_1dt to install new segment descriptors. None of the default segment descriptors in Linux specify a 16-bit default operand size.

It is convenient to use EAX/RAX in the preceding instructions, because the REX prefix for accessing RAX in the instructions used in the test happens to be interpreted as DEC EAX, which enables our test to distinguish between 64-bit mode and compatibility mode as described above by modifying the value of the register that is subsequently tested in the BT instruction. However, we need to restore the value of EAX/RAX after the mode test. One option would be to store RAX to the stack. However, that may introduce a TOCTTOU vulnerability if the untrusted code can modify the saved value. That is why we used shift operations to save and restore the original value of RAX, depending on the property that only the least-significant 32 bits of RAX are ever set at the locations where mode checks are needed.

The mode test comprises 11 bytes of instructions total. The mode test instruction sequence overwrites the value of the flags register. If the value of the flags register needs to be retained across the mode test, that can be accomplished using a matching pair of PUSHF and POPF instructions surrounding the mode test. These instructions are encoded identically in 64-bit mode and compatibility mode. It may be possible for untrusted code to overwrite the flags register value while it is saved to the stack. However, trusted code should not depend on flags register values set by untrusted code, regardless of whether that register has been loaded from stack memory or it has been set by the processor directly as a side-effect of executing instructions in untrusted code.

If the instruction sequence for testing the value of EAX/RAX used with an XRSTOR or WRPKRU instruction that is not followed by trusted code is valid in all modes that are reachable by the untrusted code, then the mode test code may be omitted prior to that value test code.

E System Flow analysis

This appendix describes the effective security policies of Intravirt to prevent information flows via system objects. The set of low-level domain specific rules combined with system object monitoring enables a simple and default information flow policy of deny all information flows from a protected object to an object outside of its domain. For example, if one domain requests mmap of a region that it does not have the capabilities for the mmap will be rejected. In the next section we detail an initial separation with two domains that puts this abstraction on display and analyzes all system calls to identify all that lead to information flows that bypass the basic domain property: i.e., by default a domain is restricted to access only those things it has capabilities for.

It is important to find all possible attack surfaces by listing up the system resources provided by the operating system. In this section we first analyze the threats and group system calls into classes. Recent work, PKU pitfall [12], describes a number of attacks toward ERIM [75] and HODOR [24], which are intra-process isolation abstractions, but this work fails to analyze the threat more systemically and does not build a complete set of security policy.

At first, the most simple and the most effective threat is to access the memory of other domains. The attacker could simply directly access any virtual address spaces in the same process, or she could use a vulnerable interfaces provided by the system or other domains. Therefore, all the interfaces which accesses the memory should be carefully designed and implemented.

In Linux, the file descriptor is shared within the process boundary. The attacker could easily scan the opened file descriptor of the victim process by scanning /proc/self/fd and the attacker is also free to control the file descriptor as well. Therefore, the attacker could simply access the content of the file or the socket, close the file descriptor and open another one with the same file descriptor number, or move the offset that the victim domain could access incorrect position of the file.

In addition, the code segment of each domain has to be protected. If not, the attacker is able to jump to the code segment at any time by simply jmp or call instructions, or by using more sophisticated attack techniques like Return Oriented Programming. Therefore the system has to provide the protection of this type of attacks, such as SFI and CFI.

Linux provides various special interfaces for exception handling, debugging and profiling. For example, signals are used for handling any special occasion of the process, and ptrace for debugging and profiling from other processes. These interfaces let the kernel shares the control of the process and provides various convenient functionalities, but it also provides the convenience to the attacker as well. The attacker could perform the memory access, code execution, and control the control flow
of the process freely with these interfaces. Therefore, Linux provides various protection mechanisms of such interfaces such as YAMA [80], but they are all bounded to the process.

Linux also provides special interfaces to access the memory of the process such as /proc/self/mem, which maps the virtual memory address of the process into a file. The attacker could simply use the well known, fully allowed file interfaces to access such an important file and access the protected memory at any time.

Lastly, while each system call has to be considered in isolation, it is important to also investigate potential threats due to concurrency leading to time of check, time of use attacks (TOCTOU). For example, an attacker may attempt to read a file while at the same time seeking to a protected region in the file. Depending on the interleaving policies enforced for the read system call may consider the file descriptor position before the seek, but the read ultimately returns data from the new seek location.

E.1 System Flow via System Calls

We systematically evaluate all the system calls in Linux to identify a set of basic potential flow and the policy. Our results are consistent with prior attacks [12] while expanding to other factors that the system calls could affect. We list up the type behaviors of the system calls and categorize them as the policy enforcement requirement.

At first, the group that the most system calls are the system calls which access resource handles such as file descriptor or shared memory. These handles, provided by kernel, are free to share in the process boundary that as mentioned in §4.4, within the process, such handles are not protected between domains and the attacker could make use of the handle to affect other domains. As well, depending on the threat model, the original resource could be important as well as the handle itself. For example, some policy could enforce only the opener of the handle could access the handle, or in some other policy, the handles which points to the special resources such as /proc/self/mem could be treated as a special handle in the policy.

The next category is the system calls which control the memory status of the process and access the memory, such as mmap(), mprotect(), and brk(). These group of system calls controls the memory map of the process, change the permission, and sometime modify the memory contents, so it is critical for the security. For example, the attacker could write code into a read/write memory page, then change the permission to executable, then jmp to the code. However, these system calls are considered to be essential for all the applications, it is not possible to simply disallow them to the users. Therefore, these type of system calls are carefully identified, analyzed, and enforced by the policy.

There is a group of system calls, which controls the process flow, such as rt_sigaction(), prctl(), and ptrace(). These system calls are able to control the flow of the process, access the memory, access to kernel, and provided various different and powerful interface with the kernel. These type of system calls have to be carefully analyzed, and the policy has be very sophisticated, because we cannot simply not allow these type of valuable, and well known system calls which are used in many applications. We will discuss signals in this paper, one of the most important example in this group.

Next, we have a group of system calls which controls the hardware, such as pkey_mprotect() and arch_prctl(). These type of system calls mainly accesses the registers or hardware specific resource, we could consider them as special system calls and we could disallow them if necessary.

The last group is the system calls which controls the privilege of the resources, such as setuid(), and chown(). These type of system calls are mainly controlled by Linux Discretionary access control mechanism and mostly requires higher privilege to be executed, we are not specifically consider this group as important category in this work.

But unfortunately, each system call does not categorized as just one group. For example, read() system call could access the file descriptor, read the file contents, and as a biproduct, it could change the file offset position. But it does access the memory as well by the read buffer input parameter. The attacker could execute read() to read a legitimate file, but contains malicious code, and use readable and executable page of other domain address for the read buffer, which will make the kernel simply overwrite the address. In the same sense, mmap() could allow file access without file related system calls by putting file descriptor in the input parameter.

As a result, each system call cannot be simply analyzed by its own behavior. We have to analyze all the possible interfaces, including behavior, input parameters, output parameters, and return values to properly derive the policy.

F Signal Virtualization with Untrusted Stack

Signals modify the execution flow of a process by pushing a signal frame onto the process stack and transferring control to the point indicated by signal handler. The primary reasons we must fully virtualize signals are because 1) Linux always resets PKRU to a semi-privileged state where domain 0 is made RW-accessible and all other domains are read-only and 2) because signals expose processor state through struct sigframe, potentially leaking sensitive state or allowing corruption of PKRU, which could lead to untrusted domain control while in the trusted domain context. As such, Intravirt must interpose on all signal delivery to minimally transition protection back to the untrusted domain mode and virtualize signal handler state to avoid leakage and corruption.

Intravirt accomplishes this by virtualizing signals so that all signal handlers are registered with Intravirt first, and second, registering signals with the kernel so that Intravirt always gains control of initial signal delivery. When a signal occurs Intravirt first copies the signal handler context info to protected memory so that the untrusted domain cannot read or corrupt it. Next Intravirt must deliver the signal to the untrusted domain, but to do so it must 1) push the signal info onto the untrusted domain stack and 2) switch the protection domain to the untrusted domain. Unfortunately, the semi-protected PKRU state does not map the untrusted domain stack as writable, so Intravirt first modifies PKRU so that it is fully in the trusted domain and then pushes the signal information onto the untrusted domain.
stack. Then Intravirt transitions to the untrusted domain mode, giving control to the handler registered in the first step.

The next challenge is that the domain switch to the trusted domain places a WRPKRU in the control path, which can be abused by the untrusted domain to launch a signal spoofing attack. By spoofing a signal, the untrusted domain could hijack the return path to its own code while setting PKRU to the trusted domain. As such, Intravirt must first add a mechanism to detect whether the signal is legitimately from the kernel or if it is from the untrusted domain. Figure 9 shows our approach that uses a special flag that resides in the trusted domain as a proof of PKRU status before WRPKRU. This flag is allocated with key 0, so it is writable only if the signal handler is invoked by the kernel which reset PKRU to default. A spoofed signal handler invocation from the untrusted domain would result in a segmentation fault that can be detected by the signal handler.

The next major issue is dealing with signals being delivered while the Intravirt system call virtualization is working in the trusted domain. This can cause bugs due to reentry, leading to potential security violations due to corrupted state. We must guarantee that our signal handler can only be invoked by the kernel once until we decide to either deliver or defer the signal and return to the corresponding state. The second problem arises out of the complex nature of adding Intravirt in between the untrusted domain and kernel, in the case where the signal is delivered during Intravirt’s handling of syscall. Unfortunately, we cannot simply ignore either these signals because that would break functionality. In this case, Intravirt must defer the signal till after the syscall is completed.

The solution to interrupted signal delivery is to emulate almost exactly like the kernel. As depicted in Figure 10, signals occurring while in the trusted domain will be deferred by adding them to an internal pending signals queue and masking that particular type of signal in the kernel. The latter step is not necessary but pushes the complexity of managing multiple signals of the same type to the kernel. Once the current operation is completed, Intravirt selects the last available signal that has not been masked by the user and delivers it.

Signals represent the most complex aspect of Intravirt. They present subtle but fundamental attack vectors while also exposing significant concurrency and compatibility issues. Intravirt appropriately handles all these cases and identifies several issues not mentioned by prior work [12].

Multithreading Design  So in the first version we had single threaded and the kernel delivers the signal to the interrupted thread. If the kernel is delivering on the backend of a syscall then we always come from domain 0 and thus the kernel can deliver to key 0 secure stack, but the problem happens if the dom 1 was interrupted and a signal comes. The kernel copies the pkru value and thus can’t pus to the dom0 stack. So it faults. Kernel couldn’t do copy to user.

To solve this we put the signal deliver into an untrusted trampoline page so that the kernel could always write and that would jump directly into IV to handle it. This worked but created the issue of a signal spoofing attack because now an untrusted domain could jmp to the untrusted stack trampoline. So we solved with a nexpertine type solution.

We have a new problem with multithreading in that this open page won’t work anymore because the page will be accessible to other threads in the same default domain. So we realize the interface provided by the kernel is just broken. To fix we modify the kernel to allow a return to both a registered stack, which is already there, and to return to a specific registered key value. So we now will always return to 0 domain and the 0 stack and never expose the data. We must then ensure that no one else registers, so we deny any registrations after initialization.

To summarize: small kernel patch to allow a default domain and deny any registration. CET also complicated the design of signal by adding another stack that must take care of during the signal delivering. We a special syscall to write to the shadow stack which allows us to push RIP to restore address and RIP to signal handler and restore token on the shadow stack so we can have the required token for switching the stack when exiting Intravirt. The similar trick is also used for virtualized sigreturn to switch to the old stack.

Multiple subdomains As we discussed, control flow and corresponding CPU state are critical to the integrity of sensi-
tive application. This applies to not only the Intravirt but also
the sandbox and the safebox. Since users can run whatever
code in the subdomains, any interruption during the execution
of boxed code can be exploited to leak data. For this reason, we
block the signal from the view of subdomains. The kernel can
still deliver signal to Intravirt signal entrypoint but we will treat
it as a signal delivered in trusted domain and pend that signal.