School of Computer Science & Engineering
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Verifying the seL4 Microkit

Verified `libmicrokit` and CapDL Mapping

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Abstract

This document reports on the formal verification of the seL4 Microkit. Specifically we report on (1) the formal specification of the Microkit library, (2) the functional correctness proof of its implementation, and (3) the verification of a mapping of the Microkit system specification (system description file, SDF) to the CapDL formalism that describes seL4 access rights. Both verification steps use *fully automated* (push-button) techniques. All artefacts are open-sourced.

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Chapter 1

Introduction

1.1 Project background

Formal verification uses mathematical methods to prove that software meets its predetermined specifications. It is one of the few techniques that can positively establish the absence of bugs in software.

The correct operation of a system depends on its underlying hardware, operating system (OS) kernel and higher-level frameworks that provide the environment in which the system executes. Consequently, bugs in the kernel or the frameworks can compromise the system's correctness and overall reliability. (Note: In this report we will use the general term "system" to refer to a user's complete, built software product at run-time; and "user" is meant in the common sense of "programmer".)

The high-performance seL4 microkernel was the first-ever general-purpose OS kernel with a formal proof of its implementation correctness, which was later extended from its source code in C to the binary code (thus taking the compiler out of the trust chain), proofs of security enforcement, and proofs of its worst-case execution time [Klein et al., 2014]. Presently, the full proof chain (including down to the binary) exists for the 32-bit Arm and 64-bit RISC-V architectures, with the implementation-correctness proof also available for the 64-bit x86 architecture. At the same time, seL4 demonstrates best-in-class performance [Mi et al., 2019], making it the ideal foundation for secure and dependable real-world systems.

However, far from being a full OS, seL4 is still only a microkernel, providing only basic mechanisms for securely multiplexing hardware [Liedtke, 1995]. Its API is policy-free and very low-level, making development of performant and correct systems on top of it costly, requiring a high level of expertise, and generally creating a high barrier to uptake.

The recently developed **seL4 Microkit** [Heiser et al., 2022] (formerly known as seL4 Core Platform) addresses this challenge by providing a small set of very simple, higher-level abstractions that are easy to use for building modular, yet still performant, systems that leverage seL4's isolation properties [Parker, 2023]. It also comes with an SDK that simplifies system generation (more on this below in the overview Section 1.2).

Microkit is still not an OS itself, but is rather a *framework* for building OS services and applications. It achieves much of its simplicity by restricting the application domain: instead of striving for generality (as the seL4 kernel does), the Microkit is designed for systems with a static architecture, i.e. ones where all components and their possible interactions (but not

necessarily the implementation of those components) are known at system-build time. These restrictions, while incompatible with desktop or cloud-hosting environments, are sufficient to support most IoT and cyberphysical systems.

The Microkit imposes an event-driven, sequential programming model on application code (also known as user code), which simplifies concurrency control and naturally leads to systems that consist of communicating sequential processes in the spirit of Hoare [1978]. This model tends to result in simpler implementations [Ousterhout, 1996]. Combined with the fact that module interfaces are enforced by verified seL4 mechanisms, that should dramatically simplify the task of verifying Microkit-based systems.

However, the full benefit of verifying such systems will be realised only if the Microkit itself is verified, as errors in the Microkit would render invalid any proofs about the behaviour of applications.

Simplicity helps here as well, as the implementation of the Microkit itself is also fairly straightforward, simple enough to experiment with a more automated verification process than what is used for the seL4 kernel itself.

Verification of the seL4 kernel used interactive theorem proving in Isabelle/HOL, which supports the construction of elaborate, machine-checked proofs. While powerful in what it can prove, this manual approach is highly labour-intensive and typically requires high expertise in Formal Methods. Furthermore, it requires manual re-verification whenever the code is changed.

In contrast, for verifying the Microkit, we turn to *automatic theorem provers*, specifically SMT solvers. These are tools that, once set up, can verify functional properties fully automatically, therefore frequently called “push-button verification”. This supports an agile, dynamic development style and should enable broader participation across the seL4 community.

Push-button verification was already deployed in the binary verification of seL4 [Sewell et al., 2013], and more recently has been used in functional-correctness proofs of simple operating systems [Nelson et al., 2017, 2019].

We note that while model checkers [Clarke et al., 2003; Holzmann, 1997] also have the “push-button” property, SMT solvers with a suitably chosen theory are particularly fitting for proofs of functional correctness.[1]

Still, significant challenges remain: The seL4 proofs were able to assume strictly sequential execution, thanks to the non-preemptible implementation of seL4 [Peters et al., 2015]. But this assumption cannot be imposed on the code that executes in user mode, such as the Microkit and any system components built on top of it, as their execution can be preempted at any time. Making verification tractable requires that this concurrency be tamed. Previous push-button verification approaches assume these challenges away, which drastically limits the guarantees that can be obtained from verification.

The Microkit deals with this concurrency challenge by observing that its functions execute atomically with respect to each other, which in turn ensures that their execution does not depend on states that can change during preemptions, except for explicitly shared memory objects (more on the notion of executing atomically in Section 5.4.4 where we introduce thread-local states).

[1] Most notably, SMT solvers are suitable for verifying Floyd-Hoare pre-/postconditions, while model checkers are failing on this score.

This report outlines our verification approach and describes the steps we have carried out so far and what gaps remain outstanding.

Specifically, we cover in detail the following, successful stages in the formal verification process for the seL4 Microkit framework:

A formal specification of the Microkit, or more specifically, of the Microkit library, `libmicrokit`, which is the interface between the Microkit and the seL4 kernel (for more detail on the structural make-up of the specification see Section 5.4).

An automated, functional correctness proof of the Microkit implementation, meaning we show that the implementations of the Microkit library plus Monitor satisfy their respective specifications (for more see Section 4.4).

A verified mapping of the Microkit system specification (written in *SDF format*) to the *CapDL* formalism [Kuz et al., 2010], which describes access rights in seL4 (as covered in Chapter 6 – 7).

We note that this mapping to CapDL allows to connect to a variety of tools for system initialisation, including a formally verified one – see discussion in Section 6.1.

For the time being there remain some gaps and limitations preventing end-to-end proofs. These are discussed further in Chapter 11.

Furthermore, an early, concise account of the work and findings of the project has been published in [Paturel et al., 2023].

1.2 seL4 Microkit Overview

1.2.1 Programming model

The programming model presented by the Microkit API is extremely simple; we only require four abstractions as introduced below and summarised in Figure 1.1. (Section 5.2 will provide additional details.)

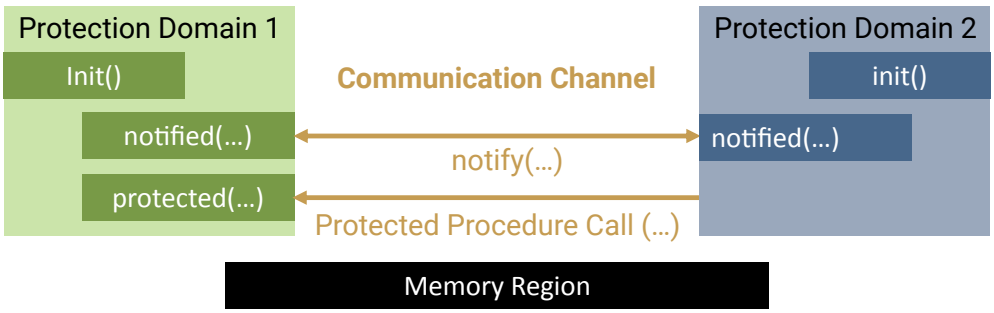


Figure 1.1: Microkit abstractions at a glance

A **protection domain (PD)** is the program abstraction, the primary abstraction of Microkit. It provides a simple, event-driven execution model, within which the user provides the implementation of a handler, called the `notified` function, to handle interrupts (IRQs) or notifications coming from other PDs. The user must also provide a (possibly empty) implementation of the `init` function, which deals with any setup specific to the PD and is called by the system exactly once at startup time, before any of the other entry points

can be invoked. The PD may also have a protected procedure `protected` to be executed when another PD invokes a protected procedure call (see below) to it.

For example, a PD can be a virtual machine that can run a Linux guest operating system (OS). Or, a PD can be a device driver, able to receive interrupt requests (IRQs) via an implicit channel.

A **communication channel (CC)** connects exactly two PDs. A PD may invoke the (non-blocking) Microkit function `microkit_notify` on a CC, which will result in the underlying system (consisting of Microkit and seL4 kernel) at some future time to invoke the `notified` function of the PD at the other end of that channel (the channel denotation implicitly identifies the notifying PD).

A **protected procedure call (PPC)** is the mechanism by the Microkit for implementing inter-PD function calls. In essence, a simple call is made to a user-provided handler `protected`. We note that the Microkit library supplies functions that allow a PD of lower priority to perform a protected procedure call but only to a PD with higher priority (see also the comment at the end of the first paragraph of Section 3.2).

A **memory region (MR)** is the final abstraction of the Microkit. It may be accessed by (i.e. mapped into) one or more PDs, with potentially different access rights, thus providing shared memory for communication.

This is all that is needed to build a complete system (i.e. a user's complete, built product at run-time – see also the note in Section 1.1, second paragraph).

A **Microkit-based system** is then a collection of concurrently executing PDs (modules) that communicate via shared memory, synchronise via notifications sent along channels, invoke servers via PPCs, and possibly receive IRQ notifications.

In particular, *systems built using the Microkit do not use seL4 kernel functions/services directly* (s.a. threads, device interrupts, etc.); *instead, they rely on above mentioned simple abstractions*, which are provided by the **Microkit's library**. Their composition is described in a formalism called the **system description file (SDF)** format, as set out in more detail in Section 5.3.

PDs are single-threaded.[2] The Microkit guarantees that the `init`, `notified` and `protected` functions execute atomically with respect to each other, eliminating the need for any concurrency control inside a PD – much in the spirit of the CSP model of Hoare [1978].

1.2.2 Implementation

The seL4 Microkit implementation consists of three components, which are linked to user-provided module implementations by the Microkit SDK:

- The `system_initialiser` is a special, privileged, system-provided PD, which allocates system resources. The design for static architectures makes it possible to restrict this allocation to system startup time.
- The `monitor` is a system-provided PD that runs after startup and acts as a fault handler, that is it receives notifications generated by the system if a user PD faults.

[2] Note that a multi-threaded PD can be constructed by having several PDs share the same address mapping and allocating them on different cores.

- The interface library `libmicrokit` is linked to each PD and maps the Microkit APIs to seL4 system calls. For each PD, it implements an event-handler loop `handler_loop` that waits for an event on any of the channels connected to the PD, and for each such event invokes the *user-provided* handler functions `notified` or `protected` as appropriate.

Figure 1.2 shows how these three components in conjunction with user PDs fit together as a whole seL4-based system.

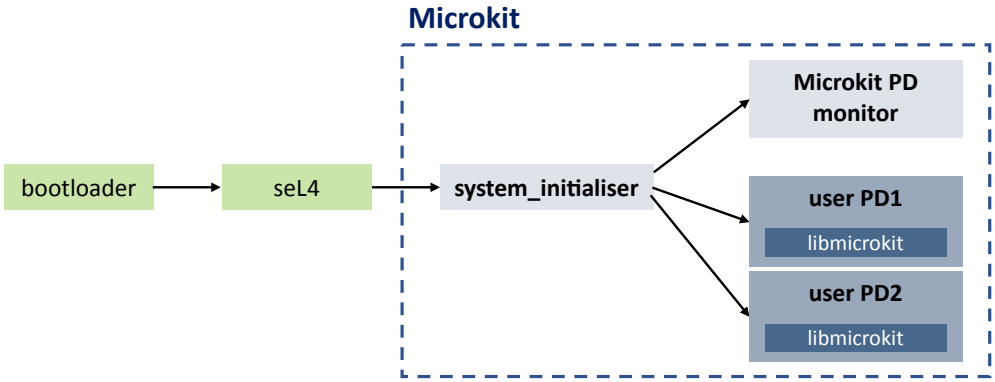


Figure 1.2: Microkit components when loaded as a whole seL4-based system.

The key Microkit-based functions can be grouped into three categories as per Figure 1.3:

1. in red the functions that are provided *by the user* (as described earlier in Section 1.2.1);
2. in blue the key functions that are provided *to the user* by the Microkit library (such as for notifying another PD or acknowledging an interrupt); and
3. in green the *purely internal* Microkit functions (most notably the event-handler loop as mentioned just above in the third bullet point).[3]

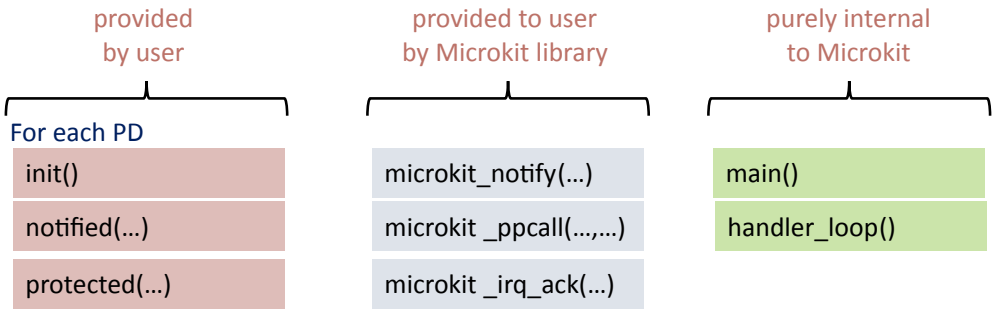


Figure 1.3: Microkit-based functions – in red: provided *by the user*; in blue: provided *to the user* by the Microkit library; in green: purely *internal* to the Microkit.

A note of practicality regarding the components: The `monitor` (see top right in Figure 1.2), while being a PD, is purely internal to the Microkit similar to other internal Microkit library functions like the `handler_loop`. Accordingly, for ease of readability, we treat from now on the Monitor as part of the Microkit library, unless explicitly stated otherwise.

Much of the research and development work is done by the Trustworthy Systems group at UNSW – find out more on their respective project webpages about the Microkit system design and about the Microkit verification.

[3]The Microkit is a very slim wrapper of seL4: the code size of `libmicrokit` is around 300 lines of C.

1.3 Summary of outcomes

The Microkit tool (also referred to as the Microkit SDK) is a command-line Python tool that processes the user-composed Microkit SDF system description file and a set of user-provided executable files in ELF format [Wikipedia, 2001], and generates a complete system image suitable for loading by the target platform bootloader (more details about this is covered in Section 5.3).

The Microkit SDK is augmented to allow for incorporating/implementing the proof steps required for verifying the Microkit and its library. At the conceptual level, Figure 1.4 gives an overview of the various constituents and steps for verifying the Microkit and Microkit-based systems.

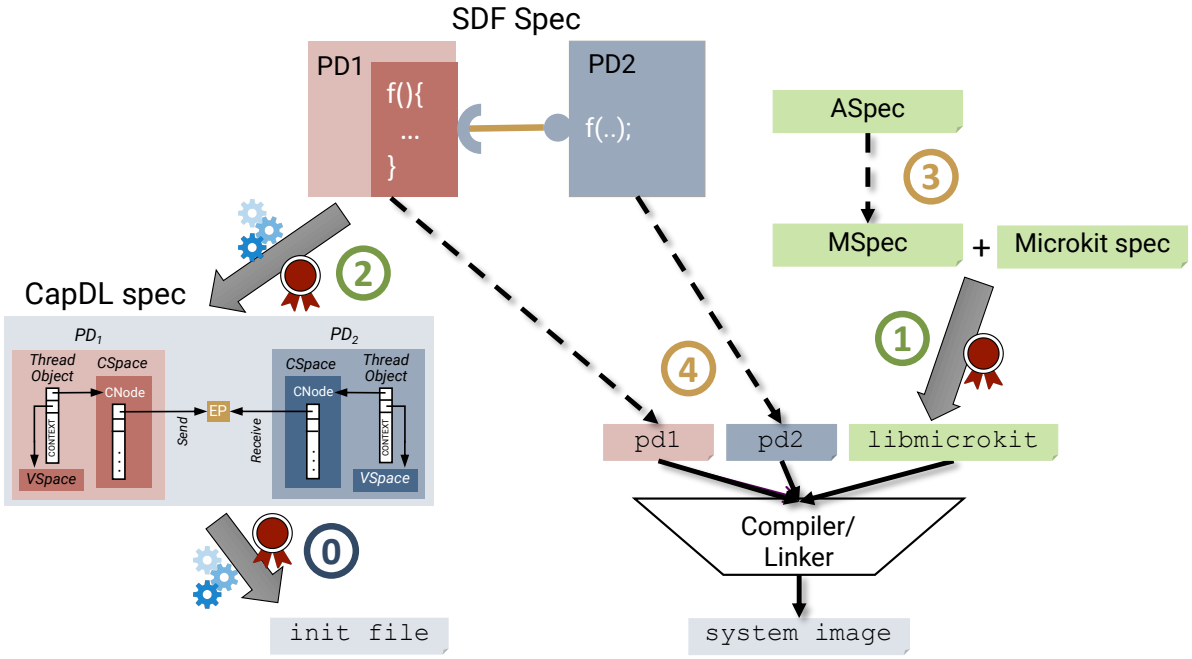


Figure 1.4: Microkit verification structure: **Proof ⓪ (system initialiser)** exists from prior work; **Proofs ① (libmicrokit)** and **② (CapDL generation)** are the result of work presented in this report; **Proof ③ (SMT-suitable abstraction)** is still required to achieve end-to-end proof (see also Section 11.2); **Proof ④** (verifying functionality of PD implementations/ Microkit-based systems) is ongoing and separate endeavour.

The seL4 microkernel has an abstract specification, the **ASpec**, against which its implementation was verified, see Section 4.1 for more details. We base the verification of the Microkit on a simplified specification, **MSpec**, which is an abstraction of ASpec. MSpec describes only the seL4 functionality required by the Microkit, and is simple enough to be usable by SMT solvers. We call this abstraction step from ASpect to MSpec the **M-abstraction**. The correctness of the M-abstraction, i.e. **Proof ③**, is not yet proved, but investigations on its verification are in progress in Trustworthy Systems (see also further discussion in Section 11.2.1).

Our eventual aim is to prove functional correctness of complete Microkit-based systems (or at least their trusted computing base) by verifying PD implementations (**Proof ④**) and composing these into proofs of system-wide properties.

For now our focus is on the underpinnings required for this ultimate aim:

Verification of Microkit-based systems depends on correct implementation of the *Microkit itself*, i.e. `libmicrokit`, (Proof ①) as well as on correct system initialisation (Proof ②), both of which are achieved by the work reported here, with the latter connecting to the verified generation of a system initialiser from a CapDL spec (Proof ⓪).

Specifically, we

1. provide a specification in Haskell of the run-time components of the Microkit (i.e. `libmicrokit` and `monitor`);
2. provide an automated verification tool called `Gordian`, written in Python (enabling the task immediately below);
3. prove functional correctness of the Microkit implementation by running our proof chain including `Gordian`;
4. provide automated generation of a specification of seL4-level access rights in the CapDL formalism (this translator is called `CapDL-tool`);[4]
5. provide a translation-validation framework that proves the equivalence of the generated CapDL from a given SDF spec.

We note that the last two items link, in principle, to earlier work on generating a verified system initialiser from CapDL, and would therefore guarantee correct initialisation of the Microkit-based system. However, as we discuss in Section 11.1, this link is currently limited to a fork of the seL4 kernel version, though still with a full proof-chain.

1.4 Synopsis of the work covered in this report

Figure 1.4 provides the context for the work covered in this report: an account of the proofs for transitions ① and ②, namely the verification of the Microkit and of CapDL generation. The early chapters 2–5 cover the verification of the run-time components, while the following chapters 6–7 get us correctly to the starting line, i.e. initialisation. The final chapters 8–13 then wrap up with presenting the tools and infrastructure used, worked examples and overall observations.

Chapter 2 on Global Correctness looks at the macro level of the Microkit correctness story, while Chapter 3 on Local Correctness explains how, as a first, crucial step towards it, we can hone in at the micro level. Chapter 4 then details the proof process we apply to verify the Microkit library and Chapter 5 discusses aspects of the formalised specification.

We then turn to the verification task of the initialisation of the Microkit when in Chapter 6 we describe the machinery we employ to verify system initialisation, CapDL, followed by a special focus in Chapter 7 on mathematical/formal aspects of the verification of the CapDL generation.

Chapter 8 details our bespoke verifier toolchain `Gordian` specifically developed as part of this project but designed to be generic enough to be able to be used in future for other push-button verification endeavours. Chapter 9 then features the various infrastructure tools that we are using for our verification tasks that already exist and been employed in a verification context. Following on in Chapter 10 we walk through the whole verification

[4] This translator emits an Isabelle model along with a C file to be used with the CapDL loader.

process with two worked examples together with their respective artefacts. In the final two chapters 11 and 12 we consider limitations, gaps and threats to validity, and following a summary of our achievements we point to the impact that this work has already had.

1.5 Repository

All artefacts discussed in this report, including this report itself together with the complete code, proof and documentation package, are open-sourced and can be accessed via the project website (under the heading “Availability”, in the latter part of that webpage).

Chapter 2

Global Correctness – The Long-Term Objective

Ultimately, our goal is to prove the functional correctness of Microkit-based systems. This requires specifying the system in terms of a *global* state machine, which holds the externally visible state of the PDs, and a *trace* (a time-ordered list of the Microkit API calls and shared memory writes made by the PDs) acting as state transitions for this (global) state machine. *Correctness* can then be defined in terms of *permissible traces*.[1]

Such a proof requires guarantees provided by the Microkit, consisting of formally verified theorems about the traces possible in any system implemented in terms of the fundamental Microkit abstractions. These guarantees in turn require a proof that the implementation of the Microkit library (`libmicrokit` including `monitor`) is correct, based on the formal specification of the kernel itself. Furthermore, the preconditions for these proofs require that `system_initialiser` correctly initialises the system.

An instructive representative of such guarantees is the following (as visualised in Figure 2.1):

Property ♣ If PD *p* sent a notification on a channel *c*, whose other end is PD *q*, using the `microkit_notify(c)` API call, then eventually (but no sooner than the next time PD *q* makes a call to receive notifications) the PD *q* will execute the `notified(c)` function.

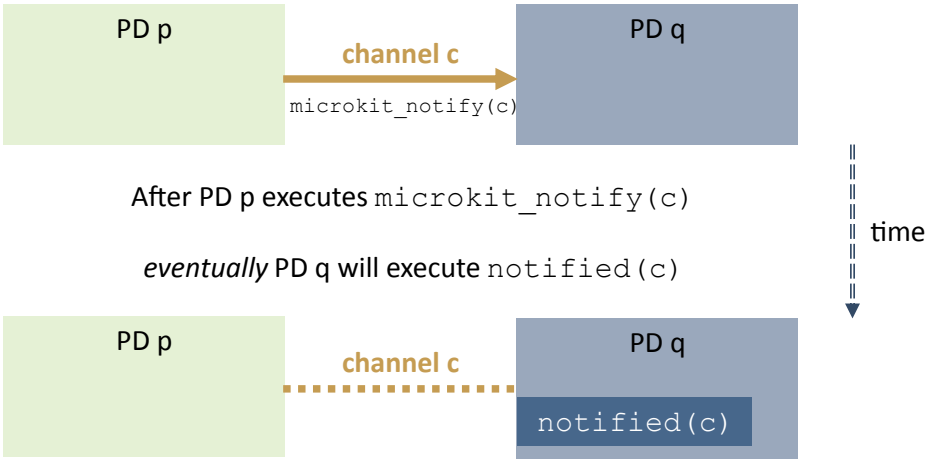


Figure 2.1: Property ♣ – a representative guarantee.

[1] Here we are considering the *traces model* for its denotational semantics.

This overarching, long-term goal requires overcoming significant obstacles, which forms part of the longer-term research vision of the project.

(a) The guarantees required of the Microkit combine complicated *liveness* (finiteness of the trace) and *safety properties* (ruling out certain finite prefixes of the trace). Verification of such properties is hard with interactive theorem proving and out-of-reach for effective automated techniques – the common theories implemented in SMT solvers do not allow reasoning about unbounded traces.

(b) One has to show that the semantics in terms of traces does not leak, in the sense that it actually captures the whole external state that a given protection domain may observe.

(c) Property ♣ can hold only subject to certain scheduling restrictions. As an example, consider PDs p, q and r, where r has a higher priority than p and monopolises q with PPCs; then, any notification from p to q will never be processed, i.e. looking at Figure 2.1, `microkit_notify(c)` will be sent by p to q but q will not be able to execute the follow-up `notified` function because of the PPCs from the higher priority r to q. While this is a legal scenario, it can be prevented by limiting r's time budget [Lyons et al., 2018], but this requires reasoning about scheduling behaviour, which is presently not defined in seL4's abstract specification, the ASpec.

Chapter 3

Local Correctness – The Task at Hand

Given the constraints as discussed in Chapter 2 above, we set about our ultimate goal of Global Correctness (proving functional correctness of a Microkit-based system as a whole) by taking a first, albeit crucial and significant step towards it: We are starting off with *Local Correctness*, where we accomplish *verification at a local level in stages* and focus on the immediate, basic work of verifying the Microkit tool itself, i.e. `libmicrokit`.[1]

Furthermore, as pointed out in the introduction, we will need to deal with verifying the system as it runs, as well as with getting to a correct starting point by verifying the initialisation of the system. The latter task we will turn to in the later chapters 6 and 7.

3.1 The local Microkit state machine

We specify the Microkit API (as implemented by `libmicrokit`) in terms of a *local state machine*, also referred to as the *local Microkit state machine*, which contains only the state pertaining to the code executing in each single PD separately, and which describes:

1. the execution state of the PD making the Microkit API call,
2. the static configuration of the user system as specified by the SDF spec, and
3. the observations that one can make during the current execution about the state of the rest of the system (such as receive calls or shared memory accesses), modeled as single-use oracles (more on this in Section 5.4.4 and Section 5.4.5).

Figure 3.1 illustrates, how closely the concepts of the local machine aligns with the model for the Microkit components as were illustrated in Figure 1.2.

3.2 Approach taken to verify the Microkit

Our formal specification of `libmicrokit` describes how the various Microkit API calls made by the currently executing PD should affect its local state. In terms of the static configuration of the system and the single-use oracles, we are able to express guarantees that a correct Microkit implementation should provide. For example, we can express that a correct

[1] Just as a reminder, in our context of Microkit verification we are usually considering the `monitor` to be part of the `libmicrokit` – as remarked earlier at the end of Section 1.2.2.

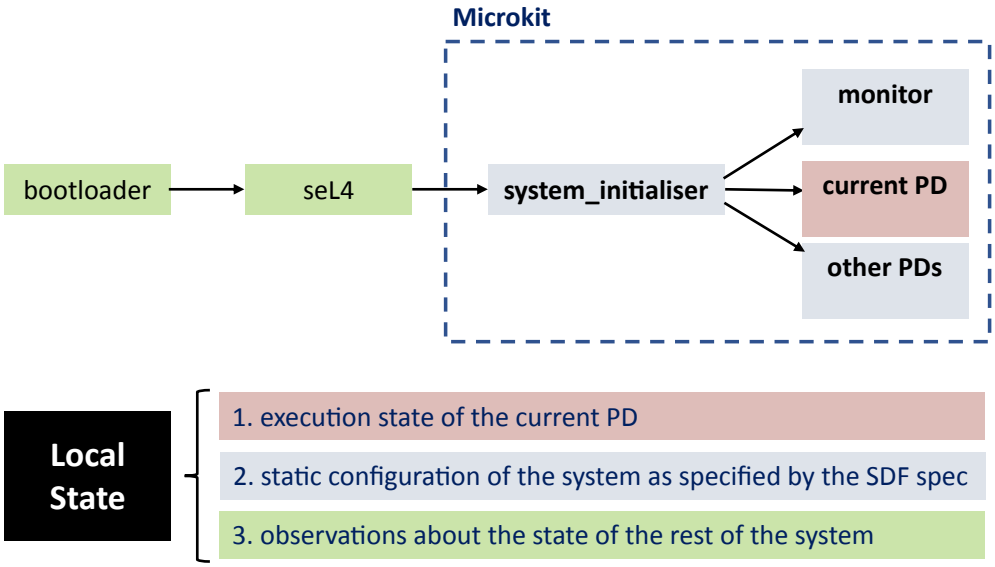


Figure 3.1: Local state capturing the three layers of the model for the Microkit components as illustrated in Figure 1.2.

implementation will not make a `microkit_notify(c)` call to a non-existent channel `c`, nor will a PPC be made to a PD that has lower or equal scheduling priority than the currently executing PD. In fact, we are able to guarantee the latter on verified systems even when the current implementation of the library does not programmatically enforce the restrictions.

As an instructive case, let us consider the correctness condition for the `handler_loop` (which is part of the `libmicrokit`), which executes on every PD:

Property ♢ The `handler_loop` never terminates. It will make a call to receive notifications and PPCs exactly once per iteration, and will correctly handle all responses returned by that receive call (according to the single-use oracle), including calling `notified(c)` if a notification was received on a channel c, and `protected(c, m)` if a PPC was received on channel c with argument m. Furthermore, it will not make "phantom" calls such as calling `notified(x)` on a channel *x* on which no notification was received.[2]

These correctness conditions were formulated with an intention behind them: Over time, we expect to make use of the facts that have been proved about the local state machine to eventually demonstrate the global correctness condition expressed in terms of traces. For example, in a global correctness argument one would appeal to the local condition Property ♢ along with a delivery guarantee coming from the kernel, to establish the global condition Property ♣ described in the previous chapter.

[2]We note that at the time of writing this report, PPCs are processed in the order of the priorities of the calling PDs while notifications are prioritised by channel ID. However, this is not inherently required and therefore might change at some time in the future.

Chapter 4

Microkit Local Correctness

4.1 Layout of formalised specification

As we progress to formalise the Microkit specification, we first give a bird's-eye view of the relevant features involved. Figure 4.1 captures the hierarchy of seL4 specs showing not only the proved refinement relations (see [Klein et al., 2014]) but also the currently not yet completed **step ③** of Figure 1.4.

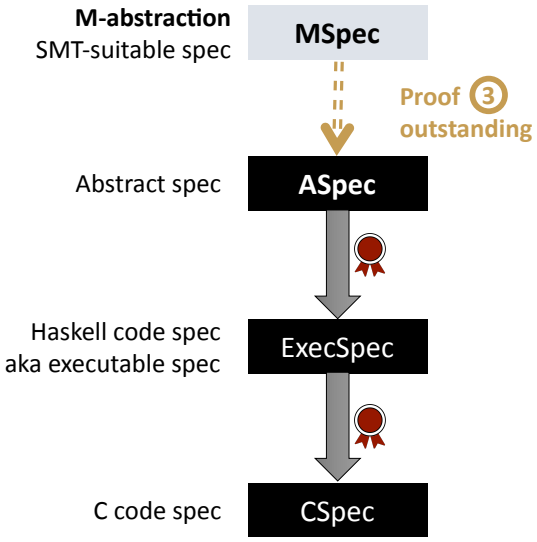


Figure 4.1: Hierarchy of seL4 specifications.

For the purpose of this report, our interest lies with Microkit-relevant specs and, hence, we will focus on **MSpec**. As mentioned earlier in Section 1.3, MSpec captures what is minimally required by the Microkit from the kernel spec. We achieve this with an appropriate (minimal/SMT-suitable) abstraction of the ASpec, which we call **M-abstraction**. To clarify: MSpec is a specification of the seL4 kernel, albeit more abstract than its ASpec, it is *not* a specification of the Microkit (as for the latter, we will simply refer to it as the Microkit spec).

For the Microkit, we are starting off writing our specification in a constrained subset of the *Haskell* programming language. The **local Microkit state machine** is then defined as an algebraic data type, and each of the Microkit APIs is defined in the form of a *weakest liberal precondition*: For each API call `f()` and property Q we describe the weakest liberal

precondition $wlp(Q)$ under which the call `f()` either does not terminate, or it terminates and upon termination successfully establishes the condition Q.[1]

The correctness guarantees are also specified. For example, the handler loop iteration is annotated with explicit pre- and postconditions ensuring that Property $\diamondsuit$ holds.

4.2 Verifying implementation against spec

We are now set up to be able to verify the `libmicrokit` implementation against its local spec. Key here is a model of the *kernel state*, and a Coupling Invariant that relates a momentary state t of the underlying kernel to the momentary state s of the local Microkit state. We give here an initial flavour of how this is achieved. Let us consider the following, fairly high-level description:

The **Coupling Invariant**[2] $s \sim t$ defines that the local Microkit state s is coupled to the current kernel state t, which can also be thought of as the current kernel state t to accurately implement/simulate the local Microkit state s.

Our definition above of the Coupling Invariant allows the **local correctness of an implementation** to be established with a *simulation-like argument*, specifically by showing the following:

- The `libmicrokit` implementation maintains the Coupling Invariant, i.e. if $s \sim t$ holds for some Microkit state s and a kernel state t, and making a Microkit API call f will leave the Microkit in some new state s', then executing the implementation of f, when starting from the kernel state t, leaves the kernel in some related state t'.
  Or, in short: for all s, t, if $s \sim t$ holds then $s' \sim t'$ holds.

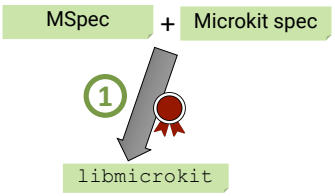
The implementation maintains all the correctness guarantees (for example, that the implemented handler loop assures the guarantee Property $\diamondsuit$ as defined above).

We note that $\sim$ allows to capture functional correctness, but will not deal with properties like liveness, i.e. finiteness of traces.

Subsequent Section 5.4, especially Section 5.4.1, goes into more detail of how to employ the Coupling Invariant $\sim$ for our local-correctness proof.

4.3 Our pipeline of verification stages

Having the overarching verification picture Figure 1.4 in mind, it is **transition step ①** that is our task at hand here.



[1] For formal reasoning purposes we allow API calls `f()` to potentially be non-terminating, however we note that all seL4 API system calls will terminate given that their latencies are all provably bounded [Sewell et al., 2017].

[2] Formerly, the term “Implementation Relation” was used.

The process of verifying the Microkit implementation proceeds via multiple steps, as visualised in Figure 4.2, and comprises for the most part automated tools or, otherwise, essentially once-off transcription/inspection-type effort. Its various features are set out in detail in Section 4.4.1 below.

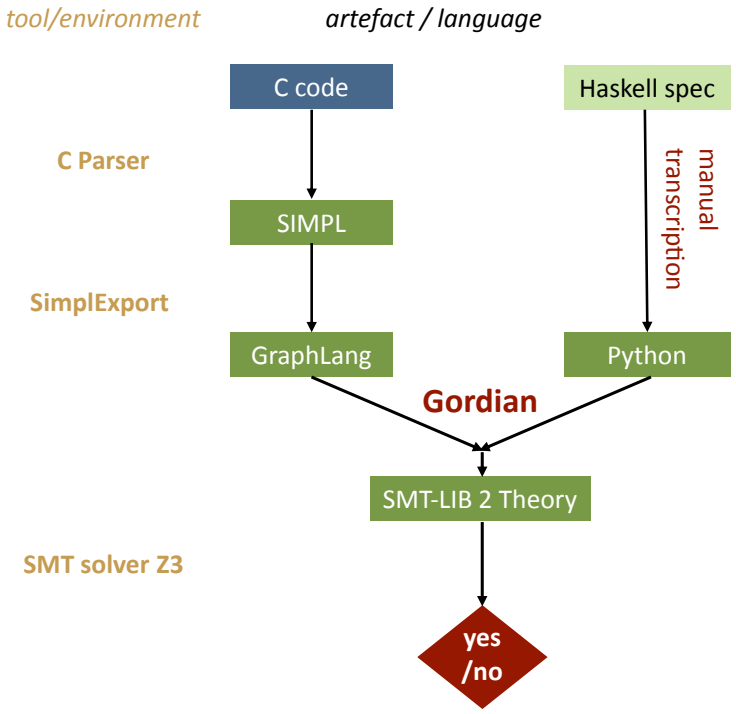


Figure 4.2: Pipeline of the Microkit verification steps – see further explanations in Section 4.4.1.

Apart from standard verification tools and environments (theorem prover **`Isabelle/HOL`**, programming language **SIMPL** and SMT solver **`Z3`**), we are reusing a number of tools and libraries developed for the seL4 kernel verification effort (see Chapter 9 for more about these infrastructure tools). Such reuse is not just for convenience. It also reduces the risk of semantic mismatch, where the assumptions of one artefact in the proof chain might not be satisfied by the guarantees of the previous one. Such mismatches can be subtle and are easily obscured by unverified correspondences.

Additionally, for the functional correctness verification task of this project we developed a bespoke toolkit called `Gordian`. It is designed to bring together the source code strand with the spec strand as an automated step – Figure 4.2 nicely displaying the central role of `Gordian`. With Chapter 8 we have dedicated a whole chapter to describe the toolkit, its components and the theoretical underpinnings.

Our verification pipeline process has been designed for verification of not only the `libmicrokit` but also to allow uplifting verification to any Microkit-based programs – key here is the general C source code that the proof pipeline and its tools accept as input together with a given set of specs of `libmicrokit` functions that `Gordian` has available to draw upon for its proofs.

We are marking out a set of *basic Microkit library functions* that any of the other Microkit functions will be composed of and with it their proofs. To ensure integrity of using these basic functions as a built-in library for `Gordian`, we need to anchor their own correctness proofs independently – more on this in Section 4.4.2.

4.4 Verifying `libmicrokit`

4.4.1 Steps of the verification process

Let us now follow Figure 4.2 and its various verification steps of the Microkit implementation. Consider a function of `libmicrokit` written in C, for example the `handler_loop`.

The C code is first processed by the same **`C Parser`** [Barthwal and Norrish, 2009] that is used in seL4 kernel verification. This tool defines a semantics for a (large but well-defined) subset of the C language and translates the C source code into a semantically-equivalent program in the **SIMPL** programming language [Schirmer, 2006]. This process guarantees that we use the same C semantics as in the kernel verification.

We then perform a semantics-preserving translation of the SIMPL code into a control-flow graph in the graph language **GraphLang**,[3] using existing **`SimplExport`** tools. Figure 4.3 gives an idea what these graphs look like, here in the case of the Microkit handler loop.

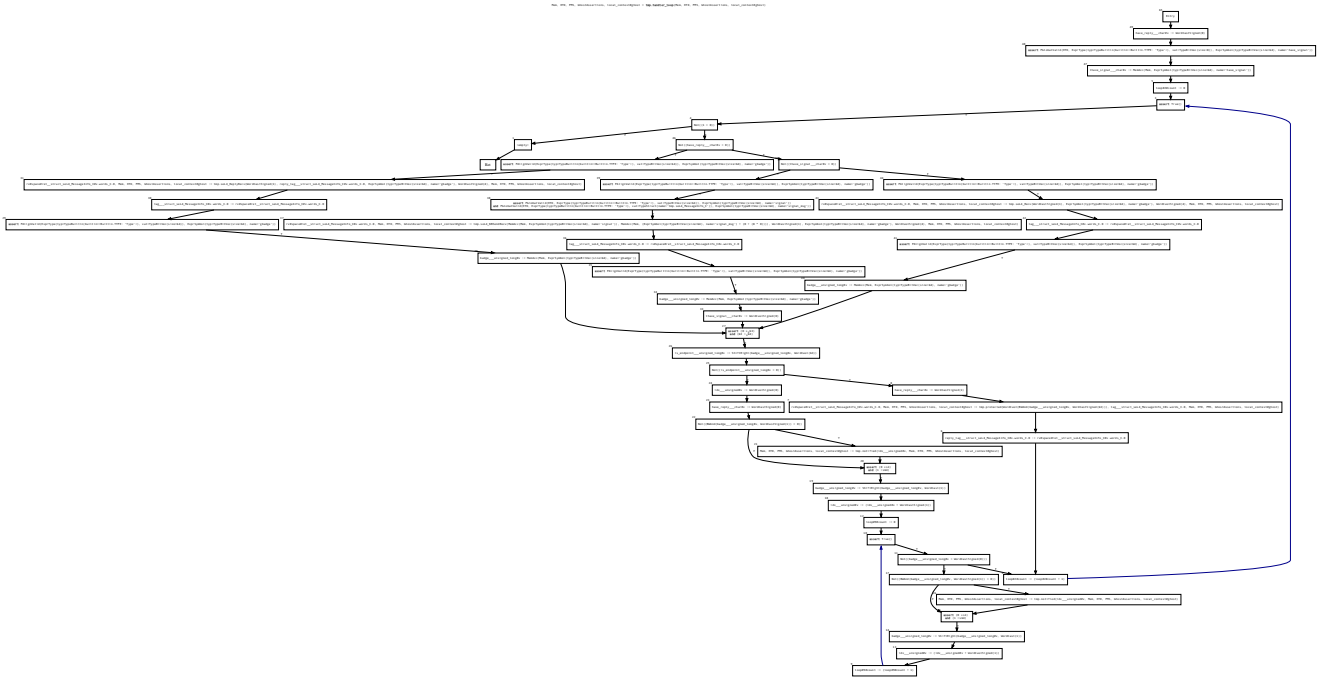


Figure 4.3: This control-flow graph is generated by the graph language tool GraphLang for the `libmicrokit` function `handler_loop`. (Here in the graph, loop edges are marked in blue, whereas `Err` nodes that result from false branches of assert statements are left out.)

GraphLang is a common intermediate language, able to represent essentially arbitrary, unstructured control-flow. It is already used in the binary verification of the seL4 kernel [Sewell et al., 2013]. (This same tool-chain will be a natural candidate for the eventual proof of the *binary* `libmicrokit`.)

Alongside an implementation of Microkit in C we have provided a formal specification for it in **Haskell**. This spec is then *manually transcribed* into **Python**.

For the next step **we have developed a new, tailor-made, automated tool,** called **`Gordian`.** It first takes the above Python description and generates a spec in **SMT-LIB 2**. `Gordian` then verifies the given GraphLang graph against the generated SMT-LIB 2 spec. It does

[3] GraphLang was formerly known as SydTV-GL, Sydney Translation Validation Graph Language.

this by annotating the graph with the function specification – Figure 4.4 is the corresponding annotated graph for the handler loop, providing an visual impression of the added complexity. From this expanded control-flow graph `Gordian` generates a

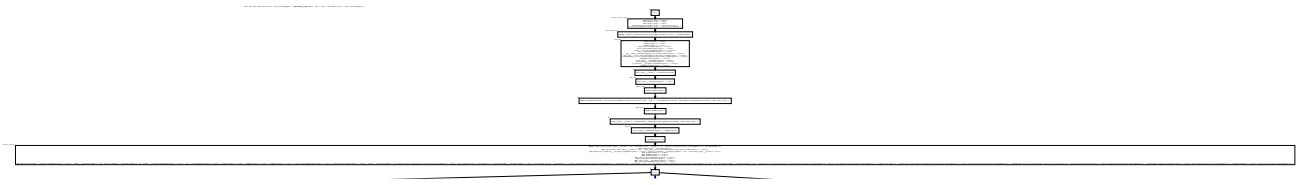


Figure 4.4: Excerpt of `handler_loop`'s control-flow graph (top right segment of Figure 4.3) when fully annotated with its function spec; the bottom box in this graph illustrates the extensive spec annotations that are added during this step of the verification process.

single, logical verification condition, using a variant of the weakest-precondition calculus for unstructured programs of [Barnett and Leino, 2005] – we will refer to this as the *Barnett-Leino–variant algorithm*. Furthermore, when `Gordian` comes across references to basic Microkit library functions it directly relies on their respective (pre-/post-condition) specs – see Section 4.4.2 how we avoid circularity and properly anchor verification of the basic Microkit library functions. Figure 4.2 highlights the key role that our `Gordian` tool plays, Chapter 8 details its particular features and capabilities.

The **verification condition** thus produced is a formula in a standard **SMT-LIB 2 theory**,[4] meaning we can employ for our reasoning the common and standardised input and output language for SMT solvers. This verification condition is then passed to the SMT solver **Z3**) [de Moura and Bjørner, 2008], which either proves the condition or provides a counter-model.

Thus, in summary, if the verification condition is proved, it means that the function's implementation in C satisfies its specification, i.e. that we have functional correctness.[5]

Amongst others, we aim in particular to ensure behavioural properties of the kind of *absence of common programming errors*, such as:

- No null pointer dereferences and inadmissible memory accesses;
- No incorrect uses of dynamic memory during program operation (no ill-typed nor dangling pointers, no out-of-bounds errors, etc.);
- No arithmetic overflows and exceptions (no signed integer overflows, no division-by-zero, no invalid bit shifts, no invalid conversions, etc.);
- No other undesired behavior (not trying to use values of uninitialised local variables, etc.).

This "good behaviour" is achieved when, as part of the semantics check, the `C Parser` generates the respective assertions, which then is fed into the subsequent functional correctness proof process.

At this point it is important to note that for our Microkit library functions, non-termination is

[4] In our case, we have mostly been using the standard SMT-LIB 2 theory QF_ABV (quantifier-free arrays with fixed size bit vectors theory), or then the standard theories for quantifiers and algebraic data types.

[5] At this stage, the weakness of our verification process here is the translation of the function spec from Haskell to Python to SMT-LIB 2: while straightforward, it is manual and not formally proved.

always an allowed behavior.[6] For example, the handler loop is specified to never terminate, and the verifier explicitly confirms[7] this property.

4.4.2 Modified approach for the basic Microkit library functions

At the end of Section 4.3 we introduced the notion of **basic Microkit library functions**. They comprise the following six functions: `microkit_notify, microkit_ppcall, microkit_irq_ack, microkit_msginfo_new, _microkit_recv` and `_microkit_ReplyRecv`, whereby the first four are known as *public* functions and the later two as *private* (or *internal*) functions.

A key characteristic of these basic Microkit library functions relates to their low level interaction with the kernel: Only these basic functions can directly make calls to the kernel. Any of the other Microkit APIs that are composed on top of these basic ones, including any of the other Microkit library functions like `handler_loop` or a user Microkit-based program, do not interact directly with the kernel[8].

As part of this project, we have verified these basic Microkit library functions using a variation of the approach outlined just above in Section 4.4.1: we use a slightly more direct, albeit more involved process, as outlined in Section 5.6.1 and motivated by the observations in Section 5.4.3 . As a reminder, `Gordian`'s full power includes relying on correctness of the basic functions. It is this feature that naturally must be turned off when submitting one of these basic Microkit library functions itself to the verification process. In its place, we are providing the SMT solver directly with the MSpec, that is that part of the formal specification of the seL4 kernel that is relevant to the Microkit. (Remember that only these basic functions can call the kernel directly and therefor only they will need to know about the spec of seL4, i.e. MSpec.) This extra effort of special verification treatment is once-off and only for the six basic functions. Thereafter, the verification process as laid out in Section 4.4.1 is applicable in the universality as intended for general Microkit-based programs and systems.

4.4.3 Relationships of specs used in the verification process

Figure 4.5 depicts how the various specifications employed in the `libmicrokit` verification process are connected, and how much or which portion of this has been accomplished as part of our project here.

Specifications – shown in Figure 4.5 as rectangular boxes

Several specs are involved as part of our verification task; they are discussed further in this report, mostly in Section 4.1–4.3 and Section 5.4.

light green/grey boxes: these specs are fully developed as part of this project:

[6] By the same token, non-termination is allowable for any program built on top of the Microkit.

[7] To prove non-termination, we set the postcondition to "false" and get the verification to succeed. Of course, we also need to ensure that the precondition is not contradictory, i.e. does not evaluate to "false" (since a false premise would always have the SMT solver succeed irrespective of the postcondition and the program). Finally, a typical way to ascertain the precondition to be non-contradictory is to provide an example of a state which satisfies the precondition.

[8] In other words, while they may *trigger* a kernel call, this will not happen as a direct call to the kernel but only via calling a basic Microkit library function

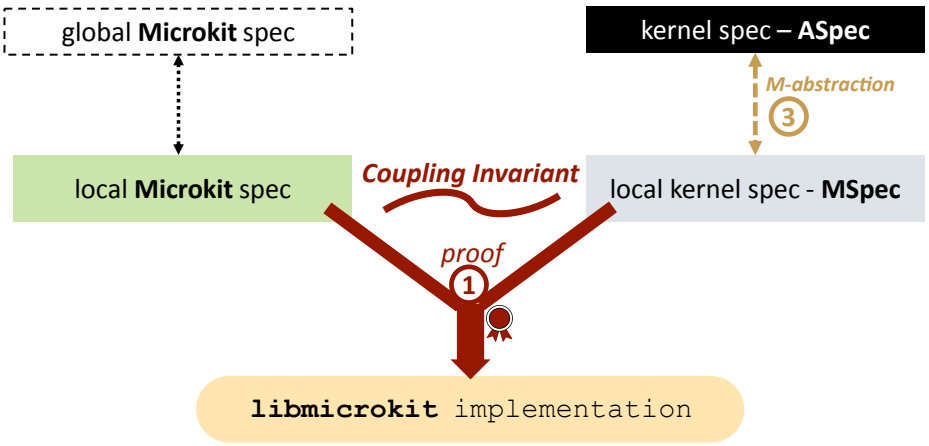


Figure 4.5: Relationships between the specifications for `libmicrokit` verification:
green ∼ and **dark red** ① are formally verified components as part of this project;
gold ③ is the verification of M-abstraction in progress;
dotted arrow is conceptualised but not formalised.

- "local Microkit spec" (usually simply referred to as" Microkit spec"),
- "local kernel spec" ("MSpec");

black box: "kernel spec" ("ASpec") is the pre-existing formalisation against which the seL4 kernel had been verified [Klein et al., 2014];

dashed white box: "global Microkit spec" is only conceptually formulated – as discussed in Chapter 2.

Linkages between specifications – shown in Figure 4.5 as arrows

Similarly for the linkages between the specifications, they are developed to various degrees as discussed across this report.

dark red arrow/relationship: the proof step ① and the "Coupling Invariant" are the core subject matter of this project and are dealt with in the wider context in Section 4.2–4.4 and in specific, great detail in Section 5.4, 5.5 and 5.6;

dashed arrows: these linkages have been conceptualised but remain unverified/unimplemented, as per discussion in Chapter 2 and comment to proof step ③ in Figure 1.4.

4.4.4 Implementation details

The GraphLang export of `libmicrokit` consists of 3,540 lines of code. The above proof chain with the `Gordian` tool is able to verify the functional correctness of `libmicrokit`. Using the `Z3` SMT solver as the main backend, the verification takes about 20 seconds on a desktop computer.

No manual proof effort is required once the spec and the loop invariants are formulated (the spec is obtained through a manual transcription from an original specification that was written in Haskell).

Chapter 5

Formalised Specification Underpinning Local Correctness

5.1 Preamble

We are now adding some more flesh to the overview spec chart Figure 4.5. The augmented Figure 5.1 now captures details about the form/kind of the various specifications, what functions are formally specified, and generally what the interdependencies are of the various specifications and semantics of the Microkit components and their seL4 kernel context when ultimately proving `libmicrokit`.

In the following we present the formalism for the specifications, the semantics and the relationships between them. We begin with the high-level spec, which also allows us to describe the general approach of verifying the Microkit. We then provide insight into formal details of the specifications to indicate how we achieve strict verification proofs.

By way of setting the scene, let us recap our overarching goal: to make seL4 easier to deploy in security-critical scenarios, while raising the assurance levels of both `libmicrokit` itself and any operating system developed using this tool kit. As mentioned in the introduction Section 1.1, unlike the *interactive* theorem-proving approach that was required for the seL4 kernel, here we have set our eyes on using the *automated* theorem-proving approach with SMT solvers. While the latter, once set up correctly, provides a sustainable assurance model (via *automatic* re-verification when maintaining underlying source code), getting to that point is still challenging for non-trivial systems, and frequently relies on strong assumptions about the environment of the code to be verified.

In contrast, here for our project we aim to assume no more than the *seL4 proof guarantees*, that is to merely rely on the results of [Klein et al., 2014], the seminal work about our verified seL4 kernel as introduced in Section 1.1.

5.2 Microkit abstractions – high-level specification

The Microkit is a minimal operating system framework with an SDK built to run on the proved secure, safe, and reliable seL4 microkernel. The kernel itself provides a small number of services for implementing systems, such as capability-based access-control primitives, device interrupts, endpoints for message passing, virtual address spaces, threads and scheduling

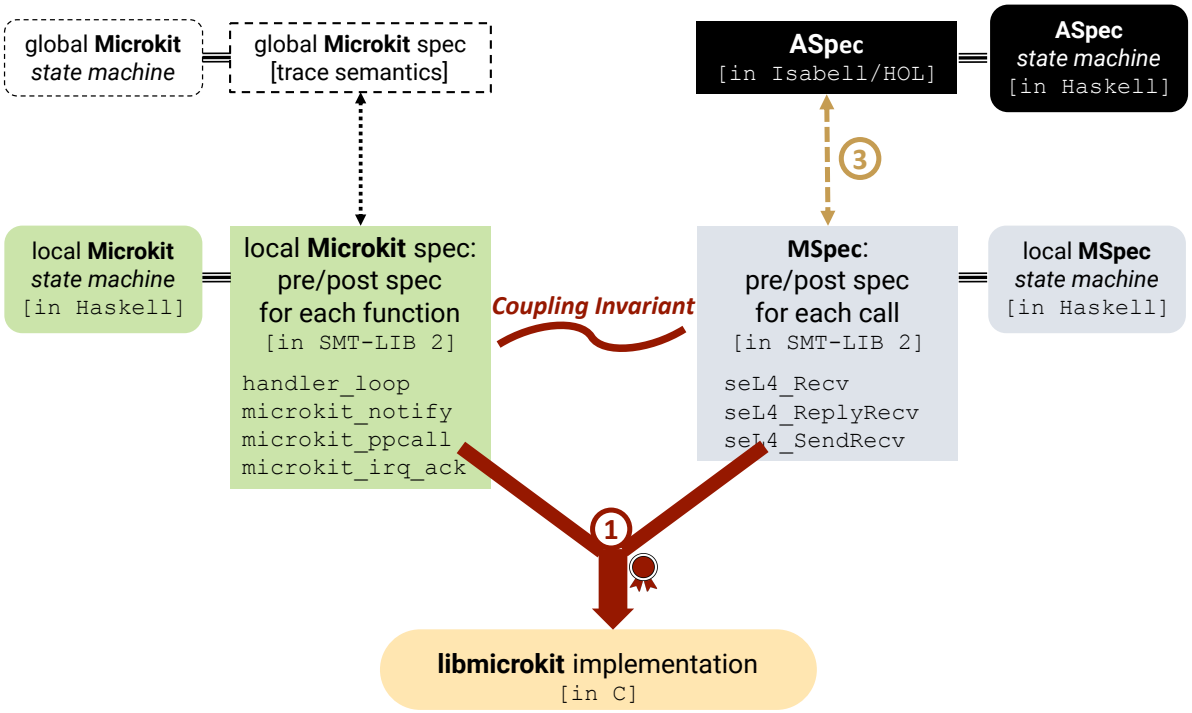


Figure 5.1: Interconnections of the specification elements for the Microkit verification.
Boxes – indicate the following: specs are drawn rectangular with their corresponding state machines shown out to the sides with small rounded corners.
Colouring – in black: prior, formally verified component of the seL4 kernel; in light green: formally verified component as part of this project; in light grey: formally specified but not (yet) verified component; in white: only conceptually, but not formally specified component.

contexts. However, *systems built using the Microkit do not use these kernel services directly. Instead, they rely on four simple abstractions provided by the Microkit's library.*

The four abstractions are PD, CC, PPC and MR, as were introduced in Section 1.2 and Figure 1.2 when we conveyed how a system implemented with the Microkit is composed of only four types of first-class objects with explicitly stated relationships between them.

These Microkit abstractions were chosen to closely mirror objects of the seL4 kernel, and were deliberately designed in such a way that they lend themselves to very efficient implementation in terms of the services provided by the kernel, while ruling out common design issues that arise when seL4 abstractions are used directly in certain inappropriate ways.

Further to their earlier definition we make the following observations about these core objects:

Protection domain (PD): The PD is the process abstraction of the Microkit: It encapsulates a thread (executing a copy of the `libmicrokit`'s event-handler loop and linked to user-provided code), a virtual address space and a capability space, all set out in a certain, well-defined manner. Each PD also has scheduling information, including a priority, associated with it.

Communication channel (CC): Two PDs may have a channel between them: `libmicrokit` provides functions for sending and receiving *asynchronous* signals across channels. On the implementation side, the existence of a channel between two PDs asserts the presence of certain capabilities in the capability spaces of the associated PDs.

Receiving and acknowledging device **interrupts (IRQs)** is also implemented using a slight variant of the channel concept, which in turn allows the design of repurposing to apply in both implementation as well as verification.

Protected procedure call (PPC): PPC is an operation on a channel that invokes the protected function in the PD across the channel, and it executes *synchronously* with the caller, i.e. the underlying system will cause the `protected` function to execute, with any return value from the `protected` function being returned to the caller before the caller can continue. The protected entrypoint is optional, meaning that only some PDs, informally called “servers”, may be invoked via a PPC – PPC is the Microkit’s equivalent of a system call in a monolithic OS. The system requires (but does not enforce programmatically) that a PPC can go only to a higher-priority PD (thus preventing deadlock).

Memory region (MR): It represents a known, contiguous range of physical memory. A memory region can be mapped into the virtual address space of (i.e. accessed by) one or more PDs with possibly different privileges. As with channels, on the implementation side the existence of a mapping between a memory region and a PD asserts the presence of certain capabilities in the capability space of the PD, and facts about the layout of the virtual address space of the PD.

In Section 7.2 we will do a deep dive and go full-out formal on specs and include an account of which capabilities are to be provided by which abstractions.

5.3 System Description File (SDF)

On the back of these abstractions we now target the formal specifications of the two main components of the Microkit, the `libmicrokit` library and the `monitor`. For that purpose, let us start with SDF, the composition mechanics of the Microkit available to the user.

A user who wants to build a Microkit-based system, first compiles and links each task, providing the `notified`, `init` and possibly `protected` functions, into a separate ELF file.

Furthermore, an XML **System Description File (SDF)** describes which PDs will be present in the system, which channels will be present between them, and which PPCs will be allowed, etc.[1] See Listing 5.1 for a simple, instructive example of an SDF specification.

These user ELF files together with the SDF completely specify the user’s system, and in conjunction with the Microkit SDK are sufficient to build the system image (using the user’s preferred build tools), as shown in Figure 5.2.

A particular feature of a Microkit-based system is its static structure in the following sense: Its properties as described in the SDF spec (such as the number of PDs, the channels between them, the mappings and permissions of the memory regions, which PDs allow PPCs, etc.) never change during the execution of the system. Therefore, in the formal model of the Microkit framework, the data specified in the SDF spec is collectively referred to as the **Microkit Invariants**. The following Section 5.4.3 highlights the practical role they play in the design of the verification approach.

[1] The description of the SDF format can be found in the Microkit manual.

```
<?xml version="1.0" encoding="UTF-8"?>
<system>
  <protection_domain name="server"
                     priority="200"
                     pp="true">
    <program_image path="server.elf" />
  </protection_domain>
  <protection_domain name="client"
                     priority="50">
    <program_image path="client.elf" />
  </protection_domain>
  <channel>
    <end pd="server" id="1" />
    <end pd="client" id="1" />
  </channel>
</system>
```

Listing 5.1: SDF specification of a simple system consisting of two PDs connected by a channel, where the PD named “server” has a PPC while the PD named “client” does not, and the channel identifier is “1”.

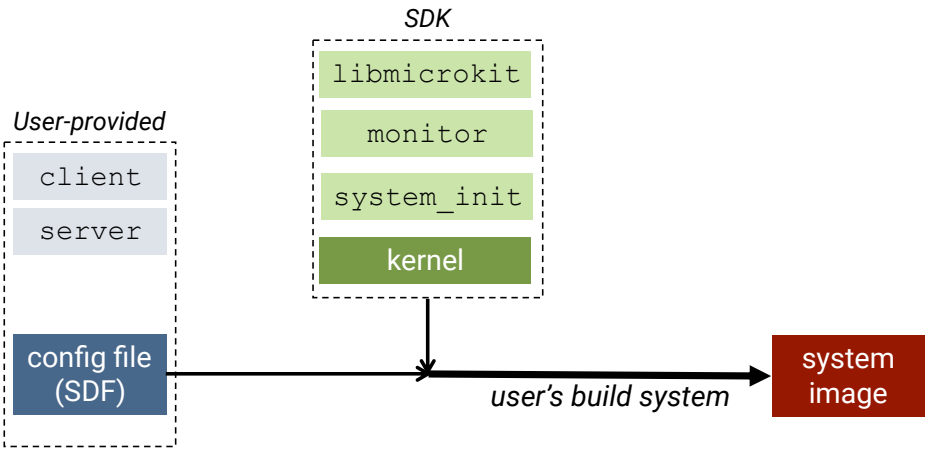


Figure 5.2: Completely built system using the Microkit: The user provides implementation of PD functions (in light grey) and a config file SDF (in darker grey); the SDK (in green) organises the Microkit library, monitor and initialisation functions plus kernel calls; the user’s own build system then constructs the final, complete system image.

5.4 Structural arrangement of the Microkit specification

5.4.1 The model: state transitions

Verifying the implementation of the Microkit library against its local specification requires several ingredients. In fact, we have already come across them in the earlier structural spec diagram Figure 5.1 but now we provide definitions and put them in context.

The main ingredients are:

Microkit State – whereby the Microkit is modelled as a state transition system, and phrased purely in terms of the four Microkit abstractions, as presented earlier in Section 1.2 and

5.2, with the corresponding spec called *local Microkit spec*.

Kernel State – whereby on the implementation side its model is again a state transition system, and which is phrased in terms of abstractions that the kernel provides.[2]

Microkit Invariants describe the static properties of a Microkit-based system (and as per last paragraph of Section 5.3 have been set in the user-provided SDF config file), namely those of: PDs (protection domains), channels, PPCs (protected procedure calls), MRs (memory regions) and IRQs (interrupts). These properties can differ between two systems, but are fixed for any one single Microkit-based system – in particular they remain fixed under Microkit-based state transitions.
Thus, not only do 'valid' Microkit states need to satisfy these invariants, but also any *Microkit state transition needs to preserve these invariants*.
Note that in formulas like the verification conditions further down, we express the Microkit Invariant properties with the simple term inv.

Microkit Dynamic State contains that portion of the Microkit state that may change with a state transition; in particular, it reflects relevant pieces of the kernel state. Thus, the Microkit Dynamic State includes parts that correspond exactly to a piece of the kernel state that is relevant to the Microkit.

Coupling Invariant is the relation[3] $\sim$ that tells us which local Microkit state x is *coupled* to which kernel state `X`, written as $x \sim$ `X`; we then say that the kernel state `X` *is consistent with* the Microkit state x.

A side remark
For our Coupling Invariant $\sim$ we are free to choose whatever we deem suitable. However:

- If $\sim$ captures only few constraints, less work is required for the proofs but we are left with only weak expressiveness. (In this case, effectively too many kernel states are related to a Microkit state, and thus, many post-program kernel states will relate to the post-transition Microkit state.)
  If $\sim$ is overreaching, it can all get too complicated and unduly detailed.
- As we will see expressed with the last clause of the verification condition, Formula 5.3e, we are looking to have the Microkit library functions with their intended correct implementation programs maintain our Coupling Invariant $\sim$.
- In short, choosing the right balance and the right kind of $\sim$ constraints is a skilled art when setting up a verification configuration.
- By way of example, Chapter 7, particularly Section 7.3 gives some insight how much detail one might want to capture within this formalism.

5.4.2 Verification condition

We can now **specify each function** of `libmicrokit` by asserting the preconditions for invoking the function, and by defining the state transition the function causes in the Microkit state – for the formulas, we capture these with the terms pre_f and $post_f$.

[2] This explains the name "kernel state" for the state of the implemented Microkit; we note that this state may also be referred to as "implementation state" or "thread state".

[3] This relation has formerly been known as "Implementation Relation"; sometimes it is also known to as "Simulation Relation" or as "Local Correspondence" .

As first introduced in Section 4.2, we can declare in standard manner:

A Microkit state transition f **is correctly implemented** by an implementation program $\texttt{F}$ precisely if whenever a Microkit state x satisfies the precondition of f and is coupled to a kernel state $\texttt{X}$, then the state $f(x)$ of the Microkit after invoking f satisfies the postcondition of f and is still coupled to the state $\texttt{F}(\texttt{X})$ of the kernel obtained by executing $\texttt{F}$ when starting from the state $\texttt{X}$.

In short (whereby $\supseteq$ indicates *is correctly implemented by*[4]):

$$\begin{array}{rcl} f & \supseteq & \texttt{F} \\ & \text{iff} & \\ \forall x, \texttt{X}.\ inv(x) \wedge pre_f(x) \wedge x \sim \texttt{X} & \implies & inv(f(x)) \wedge post_f(f(x)) \wedge f(x) \sim \texttt{F}(\texttt{X}) \end{array} \tag{5.1}$$

whereby Formula 5.1 is referred to as the *simplified verification condition* – "simplified" for the following reason:

Given that, as often is the case and particularly so with our project and its `libmicrokit` functions, the implementation programs come adorned with pre- and postconditions and can be relational in nature rather than strictly functional, we in fact work with the following composite formula as our **Verification Condition**:

$$\forall x, \texttt{X}.\ inv(x) \wedge pre_f(x) \wedge x \sim \texttt{X} \implies inv(f(x)) \wedge post_f(x, f(x)) \wedge pre_F(\texttt{X}) \wedge post_F(\texttt{X}, \texttt{F}(\texttt{X})) \wedge f(x) \sim \texttt{F}(\texttt{X}) \tag{5.2}$$

In fact, the **practical verification condition** used is written up in its simple conclusion-clause format:

For all Microkit states x and kernel states $\texttt{X}$.

$$\begin{array}{rcll} & (& & \\ inv(x) \wedge pre_f(x) & \implies & inv(f(x)) & \text{(5.3a)} \\ & \wedge & & \\ inv(x) \wedge pre_f(x) & \implies & post_f(x, f(x)) & \text{(5.3b)} \\ & \wedge & & \\ inv(x) \wedge pre_f(x) \wedge x \sim \texttt{X} & \implies & pre_{\texttt{F}}(\texttt{X}) & \text{(5.3c)} \\ & \wedge & & \\ inv(x) \wedge pre_f(x) \wedge x \sim \texttt{X} & \implies & post_{\texttt{F}}(\texttt{X}, \texttt{F}(\texttt{X})) & \text{(5.3d)} \\ & \wedge & & \\ inv(x) \wedge pre_f(x) \wedge x \sim \texttt{X} & \implies & f(x) \sim \texttt{F}(\texttt{X}) & \text{(5.3e)} \\ &) & & \end{array}$$

Putting into words, we can summarise as follows: The verification condition establishes that the Coupling Invariant $\sim$ between a Microkit state and a kernel state is preserved when accounting for satisfaction of the pre- and postconditions.

[4] We note that this notion of "is correctly implemented" reflects a sound refinement, but are aware that in itself does not deliver completness; also, it is a known cause of confusion that the usual refinement symbol "$\sqsubseteq$" points in the opposite direction to our "$\supseteq$" which accords more directly the set of behaviours.

5.4.3 Roles of the various verification condition clauses

Breaking out the verification condition into its clause form Formula 5.3a – Formula 5.3e provides a few practical insights.

Formula 5.3a – Preservation of Microkit Invariants

The first clause of the verification condition, 5.3a, simply expresses the preservation of the Microkit Invariants (as derived from the SDF spec).[5]

Formula 5.3a – Only required for basic Microkit library functions

We only need to explicitly prove Microkit Invariant preservation, i.e clause 5.3a, for the basic Microkit library functions, since such preservation is carried over to any other Microkit APIs that are composed on top of those basic ones (as for example the `handler_loop` or a user Microkit-based program).

Formula 5.3e – Limited to Microkit Dynamic State

The verification condition requires the Coupling Invariant $\sim$ to be maintained merely in terms of the kernel state being *consistent with the Microkit Dynamic State* (rather than with the full Microkit state) based on the following line of thoughts: The Coupling Invariant $\sim$ appears in the verification condition in a conclusion only in the last clause, 5.3e. Furthermore, we know from the assumptions of this clause 5.3e together with the first clause, 5.3a, that the Microkit states x and $f(x)$ satisfy the Microkit Invariants ($inv(x)$ and $inv(f(x))$) and that the kernel state $\mathtt{X}$ is consistent with the Microkit state x, in particular on the 'invariant' portion of the state x. Hence, when the requiredCoupling Invariant ($f(x) \sim \mathtt{F(X)}$) in the conclusion of clause 5.3e says "kernel state $\mathtt{F(X)}$ being consistent with the Microkit state $f(x)$" we actually need to show this merely for what is *not* invariant, i.e. we can limit the "consistency check" to the Microkit Dynamic State.

Formula 5.3c–5.3e – Kernel state references

Kernel states appear in the verification condition only in clauses 5.3c–5.3e (there they appear as $\mathtt{X}$ or $\mathtt{F(X)}$). As per Section 4.4.2 the basic Microkit library functions can make kernel calls and thus may directly access the kernel state. Therefore, any clauses in the verification condition referencing kernel states become relevant when expressing the verification condition for these basic functions.

For any of the other Microkit APIs (e.g `handler_loop` or a user Microkit-based program) quite the opposite is the case: These non-basic functions only involve Microkit states and not kernel states. Accordingly, none of the clauses 5.3c–5.3e are involved in the verification condition for those non-basic functions.

[5] Microkit Invariant preservation means that no kernel call made by any PD (which can only be effected via a Microkit library function) will ever put the system into a state that is not consistent with the Microkit Invariants. And thus, the Microkit Invariants live up to their name and will hold and no Microkit-based system implementation by a user can fail to preserve them.

Formula 5.3b – Only clause applied in `Gordian`

Bringing together above two observations “Formula 5.3a – Only required for basic Microkit library functions” and “Formula 5.3c–5.3e – Kernel state references” means that in our `Gordian` verification tool, as it is designed to be applied to non-basic functions, we *encode the verification condition effectively with the one simple clause Formula 5.3b*.

A final comment about Microkit Invariants

The verification condition also brings out that the Microkit abstractions rely on the Microkit Invariants: For any of the Microkit functions/transitions we *assume*, that is, we start off from a Microkit state x satisfying the invariant (note that $inv(x)$ only ever appears as an assumption in any of the verification condition formulas (5.1, 5.2, or 5.3a–5.3e), and from this starting position our function/transition of interest will not change that status.

The task to get to the right starting position needs yet to be tackled – see Chapter 6 and 7: As a reminder, the Microkit Invariants are described by the user-provided SDF. They constrain exactly what kernel objects the system is to have, and how capabilities to those objects are to be distributed. This relationship between kernel state and Microkit Invariants is captured and completely determined by the SDF-CapDL mapping as set out in Chapter 7. We furthermore make it a proof obligation for verified system initialisation that they initialise the Microkit from its generated CapDL description into a state consistent with the Microkit Invariants (see also the introduction to Chapter 6).

5.4.4 Thread-local State

Even though the Microkit or a user’s Microkit-based system is (part of) an operating system, it runs in user mode on top of the kernel. The kernel itself is single-threaded, but PDs run concurrently with each other and the kernel.

In principle, it would be possible to develop a Microkit state structure that represents the entire static and dynamic state of a Microkit-based system and then apply a concurrency reasoning model. However, for the moment, this would be overkill for our present purposes since we only intend to prove properties about single instances of `libmicrokit`, and each PD runs in a single thread. As part of the long-term objective of Global Correctness, dealing with concurrency will become a focal point.

Localising states

As a consequence, our formal construction of the Microkit state will not include the whole state of the Microkit, but rather only the state *pertaining to the current protection domain*. Similarly, the kernel state will not include the whole state of the system, but only the state that is *relevant to the implementation* of `libmicrokit` inside the currently running thread.

To make this precise, we introduce the notion of thread-local state. Consider the abstract specification of the seL4 kernel, the ASpec, and its abstract, global kernel state `S`, a datum d in `S` and a thread t. We then define:

The datum d is part of the **thread-local state** of the thread t if the following holds:

> In any abstract, global kernel state reachable from s, the datum d cannot change unless the thread t makes a kernel call.[6]

In particular, every sequence of kernel calls that changes the datum d includes at least one call made by the thread t.

While the relevant kernel state is a subset of the state of the entire system, it can be, and is desirably so, a suitable, small such subset; on the other hand, it is to at least comprise the thread-local state of t. Thus, we choose

$$\text{global kernel state} \supseteq \text{relevant kernel state} \supseteq \text{thread-local state of } t.$$

Analogous to the full kernel state being captured with seL4's ASpec, we refer to **MSpec** as the specification that corresponds to that Microkit suitable subset of ASpec, and the formation of MSpec from ASpec we call **M-abstraction** (see also Section 4.1 andFigure 1.4).

5.4.5 Oracle

Notification delivery and PPCs depend fundamentally on concurrency, in the sense that the received notifications and procedure calls depend on what the other protection domains chose to do. We can model the concurrency-dependent operations using *oracles*, namely variables whose values tell us what the concurrency-dependent calls *will* return in the future, when the current thread chooses to make a concurrency-dependent call.

Normally, such oracles could be represented as lists or more advanced data structures, but since our verifier encodes properties in SMT for automated verification, we had to choose a representation that avoids unbounded data structures and recursion. We opted to model the oracles as *single-use*.

Intuitively, single-use oracles are *filled* by the global state machine, and *consumed* by the local one. This allows the former to summarise and communicate complex information to the latter (notably, information only available in the traces).[7] For example, when a PD's handler loop receives an event, it consumes the receive oracle and deduces the return value from it. Beforehand, the global state machine had filled the oracle appropriately by observing the trace of the system as a whole.

More **technically**, at the beginning of each execution iteration, the oracles become available. The *availability* of the corresponding oracle becomes a *precondition* of each kernel call, so a kernel call cannot be made unless the oracle is available. Similarly, the *unavailability* of the corresponding oracle becomes a *postcondition*, so the oracle becomes invalidated, i.e. unavailable, after a kernel call which otherwise might have affected the return value of the next call.[8]

5.5 Verification condition for Microkit library

5.5.1 Microkit functional and implementation specification

With the concepts developed earlier in this section, verifying the implementation of the Microkit will collectively rely on the following:

[6] This is sometimes also known as d *is owned by* t.

[7] From the perspective of the local state machine, oracles predict the future, hence the name.

[8] The oracle becoming unavailable once used is the reason it is referred to as "single-use".

The **Microkit functional and implementation specification** consisting of

(a) pre- and postconditions for the Microkit state transitions, phrased in terms of the Microkit Invariants and the Microkit Dynamic State, and

(b) pre- and postconditions for the kernel state transitions, phrased in terms of the thread-local kernel states and MSpec.

Tying all this back to the verification condition Formula 5.2 (or, equivalently, 5.3a–5.3e), we find that those formulas, above item (a) is expressed in the terms inv, pre_f and $post_f$, while item (b) is expressed in the terms $pre_{\mathtt{F}}$ and $post_{\mathtt{F}}$.

5.5.2 The role of Microkit Dynamic State in the proof process

In observation “Formula 5.3e – Limited to Microkit Dynamic State” in Section 5.4.3 we point out that when considering the Coupling Invariant $\sim$ in our verification conditions, we can limit the Microkit states to the *Microkit Dynamic State*. For ease of communication, we will accordingly apply the prefix “Dynamic” in this context and use the term *Dynamic Coupling Invariant*.

In the context of our verification conditions we are handling the Dynamic Coupling Invariant preservation (which effectively is Formula 5.3e) by initially examining the very local (which has been carried out with this project) to then ultimately capturing the entirety of the Microkit-based system:

- Since the thread-local state t can change only when t is executing (i.e. makes a kernel call), we can reason about the parts of the implementation that involve only the thread-local state using purely sequential techniques.

  This naturally calls for a very localised obligation, namely that none of the kernel calls made by t cause the thread-local state t to cease being thread-local.

  The proof of this property relies, in a first instance, on the fact that with seL4 we use a verified kernel, and practically, it relies on not using the kernel’s full specification (ASpec), but rather using the small subset (MSpec).

- From this Coupling Invariant preservation the next step is to extend to a “thread-local preservation” (note that for our `libmicrokit` functions specifically, this thread-local preservation property indeed holds).

- Eventually, the obligation needs to deal with kernel calls that are made by *any thread* not being allowed to cause the thread-local state in MSpec to cease being thread-local.

Specification in Haskell

For practical reasons, the specifications for the Microkit and its verification conditions are written in Haskell. Such a Haskell program defines the data structures for the Microkit Invariants and the Microkit Dynamic and the kernel states, including the oracles, and it defines the pre- and postconditions for the state transitions (usually in weakest-precondition form).

5.6 Executing the verification steps

5.6.1 Proof Construction

To verify the implementation of the Microkit, we have to transform each function into a logical verification condition which can be handed to an SMT solver. This verification condition should be the most general, i.e. weakest precondition relative to the function's specification.

Thus, standard SMT solver outcome delivers our proof:

- If an SMT solver can confirm that the *negation of the verification condition is unsatisfiable*, we can consider the function *correctly* implemented.

For the basic Microkit library functions, the steps to get there are as follows:

1. The function under consideration comes with its C implementation and is annotated with its formalised specification.
2. In turn, each C function in the implementation is annotated with its formalised specification.
3. The generated verification condition Formula 5.3a–5.3e together with the MSpec (that encodes the Microkit relevant information about the proven seL4 kernel behaviour) are directly encoded in SMT-LIB 2.
4. The verification condition is checked by SMT solver.

For the non-basic Microkit function specs in general as for instance for the `handler_loop` or for a user's Microkit-based system program, the steps are as follows:

1. The function under consideration comes with its C implementation and is annotated with its formalised specification.
2. In turn, each C function in the implementation is annotated with its formalised specification.
3. The C source code is translated into SIMPL using the C Parser.
4. The SIMPL artefact is translated into the graph language GraphLang code via SimplExport.
5. As part of our `Gordian` tool, the verification condition, which effectively is only Formula 5.3b given observation "Formula 5.3b – Only clause applied in `Gordian`", is generated by the Python implementation of the Barnett-Leino-variant algorithm (as mentioned earlier in Section 4.4.1) and the invariant-finding heuristics, to then go through the verification process within `Gordian` (which calls in a SMT solver as its final step).

We note that both proof procedure steps above are set up to allow to proof check with multiple SMT solvers, which indeed has been made use of when we verified the Microkit library.

5.6.2 The Microkit properties we verify

When verifying the Microkit we verify a range of properties about the implementation of the Microkit:

1. The code in `libmicrokit` and the `monitor` does not fail, and no undefined behavior is encountered according to the C semantics induced by the `C Parser`. Among other things, this means that assertions never fail, no null pointer dereferences or out-of-bounds accesses are performed outside of user code.
2. All user-facing `libmicrokit` calls terminate, even calls with misinformed input.
3. The handler loop satisfies its specification. The handler loop never terminates. Once an iteration of the loop is completed, `libmicrokit` code will have either received all pending channel notifications and invoked the corresponding channel's user-supplied notified method exactly once, or the code will have handled a protected procedure call.

Chapter 6

Verification of System Initialisation

After having addressed the correctness of the Microkit at run-time, we now turn our attention to a *correct initialisation* of a Microkit-based system as captured by the left-hand side of Figure 1.4, specifically our **new proof step ②** combined with the subsequent, existing proof step **⓪**.

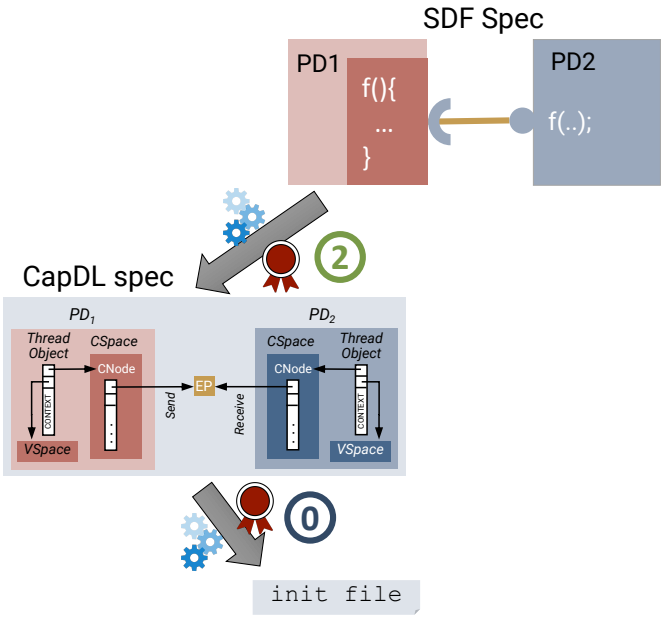


Figure 6.1: Our **new proof step ②** together with existing step **⓪** addresses the initialisation aspect of the complete proof structure overview as presented in Figure 1.4

In the prior chapter, Section 5.4.2 and Section 5.4.3, the verification condition (Formula 5.2 or 5.3a–5.3e) and the final comment convey that functional correctness of the Microkit operates amongst others on the assumption that the Microkit Invariants are satisfied; then, functional correctness delivers preservation of these invariants. From a practical perspective this assumption wants to be anchored by demonstrating that the Microkit-based system can be *initialised* into such a kernel state where these invariant properties hold – the topic of the current chapter now.

We achieve this by linking the user system spec as defined in the SDF config file with a system capability distribution, in our case defined in CapDL, and show that this link is a faithful representation. For this purpose, we can choose from a few suitable system initialisers, amongst them one formally verified one.

6.1 CapDL – the key machinery

CapDL is the language used to describe access rights in seL4-based systems [Kuz et al., 2010]. CapDL specifications can be used to track which objects and entities have access to which seL4 capabilities, and to provide complete descriptions of the capability distribution in a system running on the seL4 kernel. All this makes CapDL a powerful and versatile tool for managing seL4-based systems.

There are several tools which can initialise an seL4-based system into a state that is described by a given CapDL distribution spec. These include:

- the original `capdl-loader`, written in C and maintained by Trustworthy Systems;
- a new `rust-capdl-loader` [Spinale, 2023]; and
- a *formally verified system initialiser* called `case-init`, which is written in the CakeML language, which, in turn, has a verified compiler [Tan et al., 2016].

We augment the Microkit SDK with functionality to automatically generate CapDL language output corresponding to the system specification as captured in the SDF config file.

We note that writing the SDK in Python allows to reuse the well-tested and maintained, pre-existing Python CapDL bindings for this process. However, since it is not feasible to formally verify the functional correctness of the Python SDK directly, we perform instead a **Translation Validation**: the Coupling Invariant[1] between the *input SDF* and the *output CapDL* is shown post hoc, in each instance.

Recall that the static configuration of a Microkit-based system (i.e. the SDF spec used to generate the configuration) is one of the constituent parts of the local state machine for the Microkit Section 5.4.1. Our Coupling Invariant here then couples the static configuration of the Microkit-based system to the capability distribution in the corresponding implementation kernel state. This means that exactly one capability distribution corresponds to a valid implementation of an SDF. Note though, certain frame capabilities are not described by the Coupling Invariant yet.

With this in mind, the next step is to transcribe the Coupling Invariant suitably into Isabelle/HOL: The input SDF is imported into Isabelle/HOL as the static Microkit configuration, while the output CapDL is imported into Isabelle/HOL as the kernel state. Based on our pre-built proof schema, a simple proof script is then created in Isabelle/HOL, which can be run automated to verify that the Coupling Invariant between them, i.e. the generated capability distribution adheres to the provided SDF semantics.

In summary, given the user provided SDF the Microkit toolchain produces the following outputs:

(1) a **formal SDF**, i.e. a spec in a formal description with a clear semantics;

(2) a **CapDL description** of the seL4 system;

(3) an **Isabelle proof-script** (specific for each case, and based on the outputs (1) and (2));

(4) a **system image**.

[1] While we use the term Coupling Invariant as presented earlier in Section 5.4.1, strictly speaking here we only need a portion of that relation.

Note that it is the successful run of the proof (3) what confirms the Coupling Invariant between the user-provided SDF and its generated CapDL.[2]

6.2 Approach to formalising CapDL generation

As outlined above, the verification of the initialisation of the Microkit is based, firstly, on generating the CapDL spec from the SDF spec, and then on carrying out the Translation Validation. For the dedicated and formally/mathematically inclined reader, we set out in Chapter 7 the underlying formal definitions and a formalisation of a correct capability distribution implementation.

For now, it is sufficient to realise that the specification defines data structures for storing the Microkit state and invariants. In particular, the Microkit Invariants correspond exactly to the Abstract Systems as defined formally in Chapter 7.

> As part of our proof chain we capture the **notion of accurately reflecting the SDF to CapDL mapping in any given specific instance** with a *relation between the given Microkit Invariants and Kernel Context*.[3]

The main purpose of the relation is to express that it holds if the capabilities implied by the SDF describing the Microkit Invariants have been distributed correctly to the Kernel Context of the thread executing the current PD. This means in the formal language, amongst others, that this mapping relation faithfully manifests clauses 1–4 of Section 7.3.

As a side note, at present, the specifications are written in Haskell and then manually translated into Isabelle/HOL for the verification proof. Section 11.2.2 provides a more extensive discussion about the spec translations within the Microkit proof process.

6.3 Choosing a CapDL loader

Above sections show that for user-system initialisation we can rely on a semantics-preserving mapping from SDF spec to CapDL spec. Hence, any of our CapDL loaders mentioned above in Section 6.1 can be chosen to initialise a Microkit-based system. However, at present they all come with their drawbacks and limitations:

- The original `capdl-loader`, while versatile, is not verified and does not embed the CapDL specification in a way that is compatible with the Microkit's SDK model. We note that the `capdl-loader` must be linked against a C version of the CapDL spec, but fortunately, this can be produced from the CapDL spec itself using an unverified Haskell tool).
- The new `rust-capdl-loader` is easy to use but also not verified (see Section 9.3 for detail).

[2] We have set out in a Python file the proof to verify the CapDL generation. It contains the required definitions of the various objects like the capabilities, PDs, channels, etc, plus the respective Coupling Invariants. It furthermore includes seL4 theory files imports (e.g. for ASpec). And, finally, it lays out the required SDF-to-CapDL translation proof properties together with their (eventually very short) proof script. In Section 11.1.3 we consider its limitations in some more detail. We note that at this stage, the file has not been maintained nor re-checked against the latest Microkit verification toolchain, but is available on public Github for reference.

[3] The file detailing the mapping is available on public GitHub, but we note that at the time of writing this report it not been maintained to reflect the latest developments of the Microkit.

- The formally verified `case-init` CapDL loader is presently too restrictive, in so far as it does not support the 64-bit Arm architecture nor the new MCS variant of the kernel (on which the Microkit is based). It is also cumbersome to use, as it requires the use of the seL4 kernel build system.

Chapter 7

Formal Framework for CapDL Generation

7.1 Preamble

In this chapter we delve into the detailed mathematical formalism underpinning the specifications and proof obligations when verifying the Microkit. While neither self-sufficient nor complete, it does provide insight into how particular components of the Microkit (which interfaces to the seL4 kernel) are captured with abstract objects in a specification language. Here, we focus on the verification of System Initialisation, specifically, the Translation Validation as outlined in opening comments of Section 6.1.

As described in Section 1.2.1 and Section 5.2, a Microkit-based system is composed of a few abstract, core objects and the relationships between them. The core objects are protection domains (PD), communication channels (CC) and interrupts (IRQ), protected procedure calls (PPC), and memory regions (MR).

PDs can interact with each other expressed through various properties. For example, one such property is the presence of a channel that serves as a means of authorising communications in the user's system specification. Interactions through channels can occur in two forms: sending notifications (`microkit_notify`), which provide asynchronous signalling, and making protected procedure calls, PPCs, (`microkit_ppcall`), which allow for synchronous function calls between different PDs.

7.2 Abstract systems

In order to develop the mathematical framework to capture the Microkit's core objects with their characteristics, we begin by introducing the following definitions.

Page size k: $k \in \mathbb{N}^+$ (i.e. k is a natural number > 0);
assume k fixed for the purpose of the definitions below.

Valid protection domain identifier v: $v \in \{0, ..., 62\}$; [1]
V denotes the set thereof.

[1] In actual implementations, one bit (the 64th element) of this bit-vector is used for separate information, namely whether a notification or a PPC has occurred.

Channel identifier c**:** $c \in \{0, ..., 62\}$; [2]
C denotes the set thereof.

Endpoint (of a channel) (v, c)**:** $(v, c) \in V \times C$, where v is referred to as **end-PD**.

Channel $\{(v_1, c_1), (v_2, c_2)\}$**:** a channel is an unordered pair consisting of two **endpoints** $(v_1, c_1) \in V \times C$ and $(v_2, c_2) \in V \times C$, with:

- different end-PDs, i.e. $v_1 \neq v_2$;[3]
- at most one channel exists between two different PDs,[4] i.e. given two channels $\{(v, c_1), (w, d_1)\}$ and $\{(v, c_2), (w, d_2)\}$ then $\{(v, c_1), (w, d_1)\} = \{(v, c_2), (w, d_2)\}$, which means $c_1 = c_2$ and $d_1 = d_2$;
- an endpoint of a channel cannot be attached to two different PDs,[5] i.e. given two channels $\{(v, c), (v_1, c_1)\}$ and $\{(v, c), (v_2, c_2)\}$ then $(v_1, c_1) = (v_2, c_2)$.

Interrupts (IRQs) ***I*****:** $I \subseteq \mathbb{N}$ and I finite.
IRQ mapping (i, v, c): $(i, v, c) \in I \times V \times C$, i.e. an IRQ is mapped to an end-PD.

Memory address a**:** $a \in k\mathbb{N}$ (i.e. $a \in \{0, k, 2 * k, 3 * k, ...\}$);
A denotes the set thereof.

Permission w**:** $w \subseteq \{\mathtt{W}, \mathtt{X}, \mathtt{C}\}$, with `W`, `X` and `C` all distinct symbols (for write, execute, cached);
W denotes the set thereof.

Priority p**:** $p \in \{0, ..., 254\}$;
P denotes the set thereof.

Budget-period (x, y)**:** $(x, y) \in \mathbb{N}^+ \times \mathbb{N}^+$ and $x \leq y < 2^{64}$;
B denotes the set thereof.

Mapped memory range x comprises the following data:[6]

- a **base PD** b: $b \in V$;
- a **virtual address** a_v: $a_v \in A$;
- a **physical address** a_p: $a_p \in A$;
- a **permission** w: $w \in W$;
- a **size** n: $n \in k\mathbb{N}^+$, where $\forall m \in k\mathbb{N}$ with $m < n$: $a_v + m \in A$ and $a_p + m \in A$.

To visualise a mapped memory range including the role of size n, see Figure 7.1.

As per common usage we denote the field data for such structured objects using the *attribute* symbol "." (e.g. the base PD b of the mapped memory range x is denoted by $x.b$).

We have now all the ingredients to define the main concept:

Abstract system
It comprises the following data:

- **set of identifiers of PDs with protected procedure call** V_p: $V_p \subseteq V$;

[2] Since channel identifiers mirror valid protection domain identifiers, we also have only 63 elements available.
[3] We note that self-channels are not required for the Microkit (in fact, they are not allowed), thus greatly simplifying concurrency handling down the track.
[4] While theoretically not strictly required, for performance reasons this property is desirable and thus worth while explicitly proving and make available; also, note that it can readily be captured, and typically is so, as part of the Microkit Invariants.
[5] This property can readily be captured, and typically is so, as part of the Microkit Invariants.
[6] For demonstration purposes here we choose the concept of "range" over "region" purely for brevity and clarity of the abstract definitions; modifying for "region" requires only superficial adjustments.

The key parameters are fixed as follows:
page size **k** = 4Ki (we drop "Ki" from now on)
addresses **A** = {0, 4, 8, 12, 16, 20, 24, 28}
virtual address $\mathbf{a_v} = 4$
physical address $\mathbf{a_p} = 16$
size **n** = **8**

The requirement on the size n is:

m=0: $a_v+m = 4+0 = 4 \in A$
$a_p+m = 16+0 = 16 \in A$
m=4: $a_v+m = 4+4 = 8 \in A$
$a_p+m = 16+4 = 20 \in A$

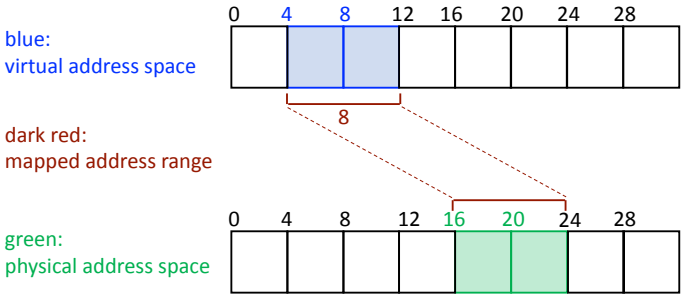


Figure 7.1: Example of a mapped memory range.

- **set of channels** C_h (this is a subset of all unordered pairs of end-PDs, and loosely expressed as $C_h \subseteq \{(V \times C), (V \times C)\}$);
- **set of IRQ mappings** I_m: $I_m \subseteq I \times V \times C$ (i.e. a set of IRQs mapped to end-PDs);
- **finite set of mapped memory ranges** M;
- **priority map** $p_r : V \to P$;
- **budget-period map** $b_p : V \to B$.

It is subject to the following conditions:

(i) **All channels are disjoint**, i.e. if $r, s \in C_h$, then either $r = s$ or $r \cap s = \emptyset$.

(ii) **IRQ ends and channel ends are disjoint**,
i.e. if $(i, e) \in I_m$ with $e = (v, c)$, and $r \in C_h$ with $r = \{(v_1, c_1), (v_2, c_2)\}$,
then $e \notin r$, i.e. $(v, c) \neq (v_1, c_1)$ and $(v, c) \neq (v_2, c_2)$.

(iii) **Virtual addresses (of mapped memory ranges) are disjoint**,
i.e. for any two mapped memory ranges $x, y \in M$ with the same base PD $x.b = y.b$,
then either $x.a_v + x.n \leq y.a_v$
or else $y.a_v + y.n \leq x.a_v$.

7.3 Accurate implementation with a capability distribution

With all the machinery in place, we can now formally capture the notion that an seL4 kernel state implements an abstract system. This in turn underpins the formalism as sought in Section 6.2.

Take an abstract system as defined above in Section 7.2. Consider a finite set F of user ELF files and an assignment from PDs to these ELF files, $f : V \to F$.

We then define that (the capability distribution of) a kernel state **accurately implements** a given abstract system if the following conditions hold:

1. For each PD $v \in V$, the file $f(v)$ is linked against the `libmicrokit` library and has that library's `main()` function as its entry point.
2. There is a PD *Monitor* with an allocated thread control block (TCB) object in physical memory, referred to as the *MonitorTCB*, along with a synchronous IPC endpoint, referred to as the *MonitorEndpoint*, and a reply object, referred to as the *MonitorReply*. The MonitorTCB is not suspended, and has priority 254 and max-priority 254. The Monitor has a single `CNode` as its `CSpace`, which contains capabilities to MonitorReply in

slot 4 and to MonitorEndpoint in slot 74. The `VSpace` of the Monitor consists only of the frames required to contain the Microkit monitor executable.

3. For each IRQ $i \in I$ there is an IRQ handler capability that allows a thread possessing it, firstly, to set an endpoint which will be notified of the incoming interrupt i, and, secondly, to acknowledge received interrupts of the same number i.

4. For each PD $v \in V$, there is a unique TCB object, referred to as the vTCB, along with a unique endpoint, the vEndpoint, and a unique IPC buffer' (used to transfer PPC arguments). The vTCB is not suspended, and has priority $p_r(v)$ and max-priority $p_r(v)$. There is a unique scheduling context, referred to as the vSC, which has budget-period $b_p(v)$. The Fault endpoint of v is the MonitorEndpoint. There is a unique notification object, referred to as the vNotificationObject, which is bound to the TCB object. (This allows the PD to receive both PPC invocations and notifications on vEndpoint; by convention, the most significant bit of the badge then allows the receiver to distinguish between the two.)

   The `CSpace` of the PD v consists of one `CNode`, which contains the following capabilities:

   - An unbadged RW capability to the vNotificationObject in slot 1, referred to as the vInputCap.
   - An unbadged capability to the `VSpace` of v in slot 3.
   - An unbadged capability to the v reply object in slot 4, referred to as the vReplyCap.
   - For each $u \in V$ and $c, d \in C$ with $\{(v, c), (u, d)\} \in C_h$: an RW capability to the uNotificationObject in slot $10 + c$, badged with 2^d.
   - For each $u \in V_p$ and $c, d \in C$ with $p_r(u) > p_r(v)$ and $\{(v, c), (u, d)\} \in C_h$: an RW capability to the uNotificationObject in slot $74 + c$, badged with $2^{63} + d$.
   - For each $i \in I$ and $c \in C$ with $(i, v, c) \in I_m$: an unbadged minted copy of the IRQ handler capability for i in slot $138 + c$.

   The `VSpace` of the PD v contains mappings for the frames required to contain the user ELF file $f(v)$. For each mapped memory range $m \in M$ with $m.b = v$, a frame is mapped with corresponding physical and virtual addresses, write permission if `W` $\in m.w$, executable permission if `X` $\in m.w$, and cached if `C` $\in m.w$.

7.4 Abstract system as used for the `libmicrokit` proofs

The framework of an abstract system is the foundation of the correctness proofs of the `libmicrokit` as carried out for this project. The central file with specs and proof script is written in standard SMT-LIB 2 syntax allowing to directly feed it to the SMT solver `Z3` and thus facilitating the automated proof process.

The account starts with the spec for an abstract system involving a Microkit state and a kernel state. Together with this, memory region (in the file defined as the datatype `MMR`), well-formedness (`wf_MicrokitInvariants`), capability mapping (`relation_cap_map`) and eventually a Coupling Invariant (`relation`) are defined, as well as pre- and postconditions for Microkit and kernel state transitions. With this arrangement the libmicrokit implementations (in terms of seL4 kernel functions) are then proofed correct by the SMT solver `Z3`.

For example in the case of `microkit_notify`, which is implemented with the kernel function `seL4_Signal`, the proof is set up to first show that notify's precondition implies signal's precondition and then that signal's post condition implies notify's post condition. Similarly for

the other functions `microkit_ppcall`, `microkit_irq_ack`, `_microkit_recv` and `_microkit_ReplyRecv`. (We note that `microkit_msginfo_new` is not treated in this fie since its implementation is a straight inline call to `seL4_MessageInfo_new`.)

Chapter 8

Our New Verification Tool `Gordian`

8.1 The process steps

The `Gordian` verifier is a newly developed, automated tool designed for the verification of functional correctness of the Microkit specifically and C programs more generally. Section 4.4 frames the context for which `Gordian` was developed with Figure 4.2 highlighting the central role it plays.

Before the `Gordian` verifier can be invoked, the target C source code must first be exported to the graph language GraphLang. This conversion is performed by the `C Parser`, the same C-to-Isabelle parser that is used in the verification of the seL4 kernel itself [Barthwal and Norrish, 2009]. The C Parser reads the C input and emits code in Norbert Schirmer's SIMPL language [Schirmer, 2006], which, unlike C, comes equipped with a formal (big and small-step) operational semantics.

The SIMPL code is then converted to GraphLang using `SimplExport`, which is a verified tool originally implemented (and still used) as part of the translation validation of the seL4 kernel. The resulting GraphLang code is an unstructured graph program representing a control-flow graph, whose semantics refines the C semantics induced by the C Parser. The represented graph consists of three sorts of nodes: *Basic* nodes which update local variables; *Cond* nodes which perform conditional jumps; and *Call* nodes which perform calls to other functions. See Figure 4.3 and Figure 10.1 as examples of such a graph.

At this stage, pre- and postconditions are being inserted in the graph: These conditions are picked up by `Gordian` from a submitted spec file (written in Python) and manipulated and incorporated into the control-flow graph as the additional spec information. In this manner we express that before a function call one needs to *assert its precondition* and after a function call one gets to *assume the postcondition*. Figure 4.4 or Figure 10.2 are the respective examples of these expanded graphs.

The `Gordian` verifier then processes its input in GraphLang along with a specification written in a combination of Python and SMT-LIB 2. Based on these inputs, `Gordian` generates a logical verification condition using a variation of Barnett and Leino's weakest precondition calculus for unstructured programs (see Section 8.2 for some technical details).

Finally, `Gordian` passes the verification condition to an SMT solver (for this project, `Z3` has been chosen), which tries to verify its validity. The successful verification of the logical formula indicates that the C function under consideration conforms to its specification. If the

verification fails, the SMT solver produces a counter-model.

We note that `Gordian` generates its verification condition in SMT-LIB 2, a common and standardised input and output language for SMT solvers. To a large extent, the logic theory used is the quantifier-free theory of arrays and bit-vectors, QF_ABV, the same, standard theory used for the binary verification proof of the seL4 kernel, and one which is supported in all contemporary SMT solvers. In some cases, we are resorting to the more flexible, but still standard theories for quantifiers and algebraic data types. In practice, using SMT-LIB 2 allows us to increase both performance and certainty by running the verification condition through multiple SMT solvers.

8.2 The verification condition algorithm

We use a minor variant of the "weakest precondition for unstructured programs" algorithm of Barnett and Leino to determine the weakest conditions that need to be met in our control-flow graphs to prove correctness:

- We first apply a *loop-elimination transformation* to the GraphLang graph, resulting in a control-flow graph that is free from cycles and that “traces correctness”, whereby the latter means that this modified graph represents a progam that is correct only if the original program was correct. (This means that if the modified graph is shown correct, we then know that the original graph must also be correct; however, if we obtain a negative result, we are not quite sure what it means for the original graph.)
- The next step involves transforming the program into a *control-flow graph in dynamic single-assignment form*. This change makes it easier for SMT solvers to do their job. In `Gordian` we have implemented the loop elimination and the dynamic single-assignment transformation, along with a comprehensive test suite.
- The final step applies the *weakest precondition computation* to the acyclic single-assignment control-flow graph to generate the verification condition to be handed to the SMT solvers.

Chapter 9

Existing Infrastructure Tools for Verifier Implementation

9.1 C Parser

In our toolchain, the same C Parser tool, `C Parser`, has been utilised that is used to create the Isabelle/HOL semantic model of the seL4 kernel [Barthwal and Norrish, 2009]. This tool translates the C source code into Schirmer's SIMPL programming language, which has a well-defined operational semantics in Isabelle/HOL. This, in turn, allows the same semantics to be used here for the Microkit as that for the seL4 kernel verification. Additionally, SIMPL can be translated into the graph language GraphLang using the `SimplExport` tools in a way that preserves the semantics.

The `c-parser` handles a reasonable subset of the C99 standard, and accounts for architecture-defined behaviours such as endianness, or the number of bits in an `int`. The most important limitations of the subset of C implemented by our Microkit verification toolchain are the following:

1. No `goto` statements
2. No fall-through cases in `switch` statements
3. No type unions
4. No taking the address of stack variables ("automatics" in C).

These operations are already absent from the Microkit implementation, specifically the Microkit library, and adherence to these restrictions are not expected to become a constraint or difficulty in future development of the Microkit.

We note that an early version of the source code of the `handler_loop` in `libmicrokit` did contain one instance of the fourth limitation, where the address of the local variable `badge` had to be taken as a result of a peculiar design decision in `libsel4`. It turned out that a simple workaround was available by introducing an auxiliary global variable to store this badge (required changes: 4 LOC).

Due to the fact that the Microkit library and monitor are implemented in terms of other libraries such `libsel4`, which were not written with the C Parser in mind, we had to write preprocessing scripts before we could read the `libmicrokit` and `monitor` using the parser and successfully pump it through to obtain an artefact in the GraphLang language using

`SimplExport`. Options to automate this process in a way that is resilient to Microkit implementation changes are being explored.

Statistics

Taking all includes together, exporting `libmicrokit` yields a graph consisting of 5,627 nodes, divided into 320 functions. However, 302 of these come from `libsel4` and other includes (of which only 6 are relevant to the verification, and which have been already covered with the `libsel4` verification process), leaving a total of 18 functions (about 110 nodes) to be verified.

The main `handler_loop` consists of 42 nodes, the `monitor` of 46 nodes.

9.2 GraphLang and SimplExport

The GraphLang language (formerly known as SydTV-GL, Sydney Translation Validation Graph Language), being used to describe control-flow graphs was originally developed for, and forms an integral part of, the seL4 kernel's translation validation and worst-case execution time analysis proofs.

Our `Gordian` verifier also uses GraphLang as its foundation. As per Figure 4.2 and Section 4.4.1, in one of our verification process steps we employ the `SimplExport` tool to obtain the intermediate artefact written in this graph language. This approach allows us to create an implementation that is compatible with previous advancements and to reuse existing code for parsing GraphLang. We improved the type-safety of the existing GraphLang toolchain by adding new interfaces with additional checks and constraints. These fine-tunings not only help to prevent programming errors in the new developments, but can be upstreamed to increase the overall trustworthiness of our toolchain.

The external translation validation tools in connection with GraphLang are undergoing development as part of the *Binary Correctness and Multicore Verification* topic area. Once concluded and stable, we expect resulting improvements to be suitable to be incorporated into the translation validation version of our toolchain `Gordian`.

9.3 Rust CapDL loader

Nick Spinale has recently developed a new CapDL loader implemented in Rust [Spinale, 2023], accompanied by a simple, self-contained tool that serialises a CapDL spec and adds it to the pre-compiled loader binary. This scheme aligns it well with the Microkit ecosystem. This Rust CapDL loader also supports a configuration whose operation is more akin to that of the unverified C loader today, allowing for a simpler and space-efficient loader, as there is no deserialisation at runtime. In fact, under this configuration, the loader does not even require a heap allocator. For our purpose though, it is noteworthy that to date, this Rust CapDL loader is not as yet verified.

Importantly and in summary, our verified version of the Microkit does, along with the verified CASE initialiser, support (and is integrated with) Spinale's new Rust CapDL loader.

Chapter 10

Worked Examples – Walk-Through Of The Verification Steps With Artefacts

10.1 Arithmetic sum `arith_sum`

Summary description

$s = arith\_sum(n) =$ sum of all the natural numbers less than the given natural number n

Mathematical spec

precondition: $0 <= n \ \wedge \ n <= 100$
loop invariant: $0 <= i \ \wedge \ i <= n \ \wedge \ s = (i-1) * i/2$

(Note that an upper limit, in our case "100", ensures avoiding overflow.)

C code (from the repo for demo examples)

```
1  // pre condition (0 <= n <= 100)
2    int arith_sum(int n)
3  {
4      int s = 0;
5      for (int i = 0; i < n; i++)
6      // loop invariant (0 <= i <= n && s == (i-1)*i/2)
7      {
8          s += i;
9      }
10     return s;
11   }
```

Listing 10.1: C code for the Arithmetic Sum function `arith_sum`

Verification stages/steps

1. Converting into GraphLang => *control-flow* graph Figure 10.1
2. Adding the function spec and turning graph into dynamic single-assignment form and acyclic $\Longrightarrow$ expanded control-flow graph Figure 10.2
3. Computing Weakest Precondition (per Section 8.2) $\Longrightarrow$ Logic formula in QF_ABV

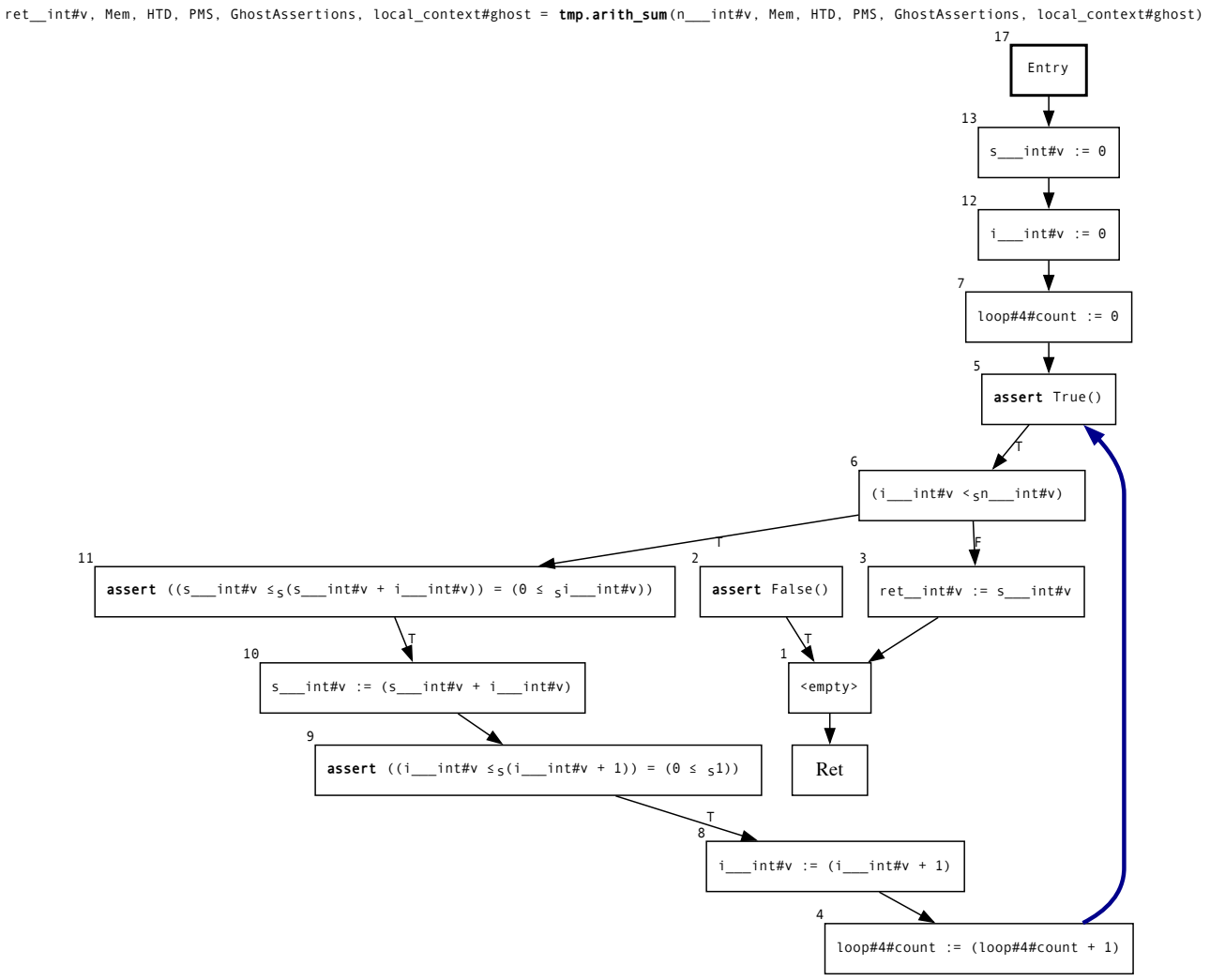


Figure 10.1: Control-flow graph for Arithmetic Sum, as generated from SimplExport.

4. Incorporating manually[1] the invariant to SMT tool
5. Running SMT solver Z3 $\Longrightarrow$ Result = “Yes”

[1]This step is in the process of being automated.

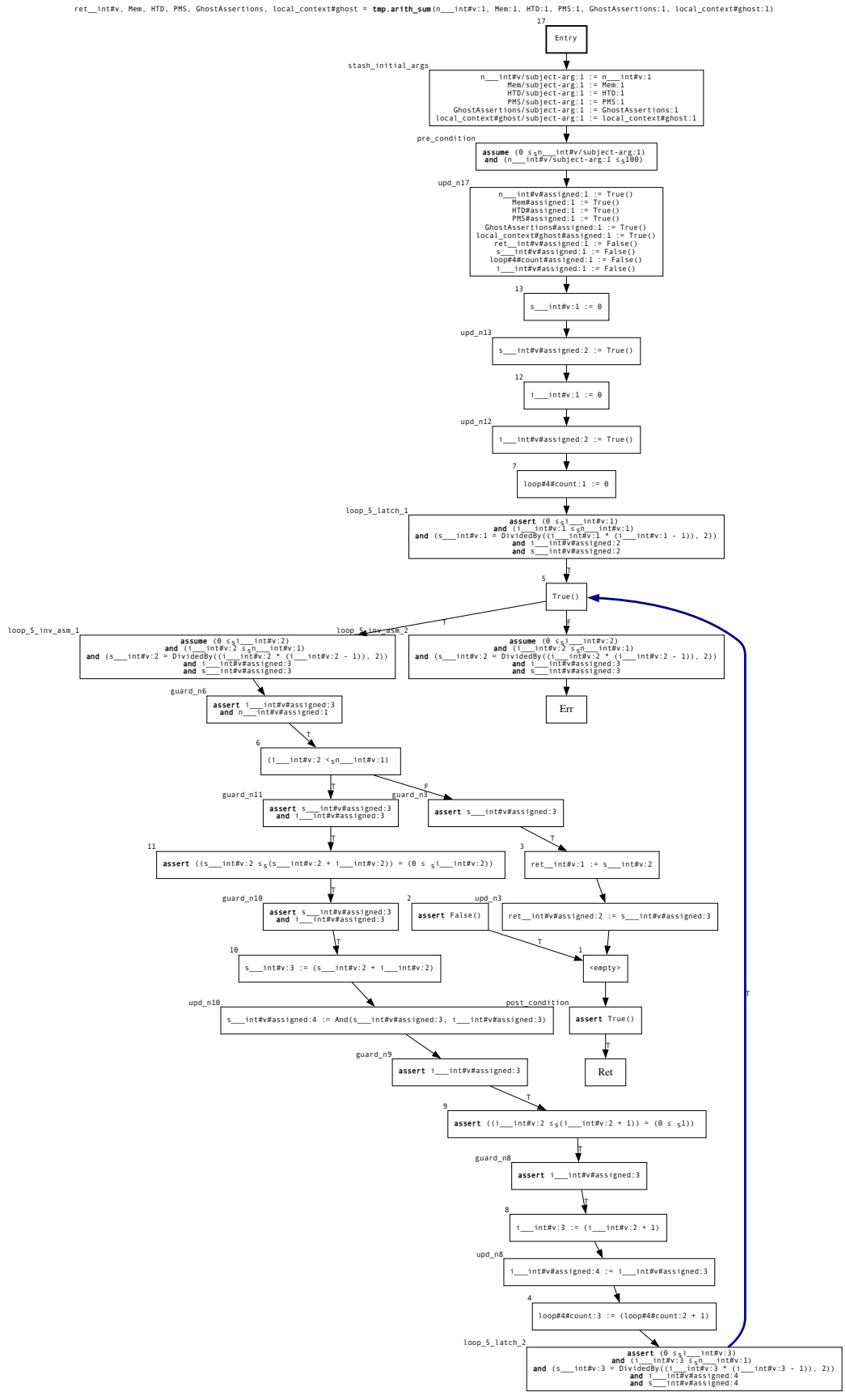


Figure 10.2: Control-flow graph for Arithmetic Sum, with spec and in dynamic single-assignment form and acyclic.

10.2 Microkit Monitor `monitor`

Summary description

The monitor's purpose is to act as the fault handler for its dedicated PD (that is, each PD has its very own monitor).

Specification

The monitor is to exhibit "NO undefined behaviour" in the sense of "absence of common programming errors".

C code (the code here is a slightly sanitised version from our repo for the example `monitor`)

```
monitor()
{
    for (;;) {
        char cont = 0;
        seL4_Word badge, label;
        seL4_MessageInfo_t tag;
        seL4_Error err;
        tag = seL4_Recv(fault_ep, &gbadge, reply);
        badge = gbadge;
        label = seL4_MessageInfo_get_label(tag);
        seL4_Word tcb_cap = tcbs[badge];
        if (label == seL4_Fault_NullFault && badge < 64) {
            err = seL4_SchedContext_UnbindObject(scheduling_contexts[badge],
                    tcb_cap);
            err = seL4_SchedContext_Bind(scheduling_contexts[badge],
                    notification_caps[badge]);
            cont = 1;
        }
        if (cont==0) {
            if (badge < 64 && pd_names[badge][0] != 0) {
                puts(STRINGLITERAL);
                puts(pd_names[badge]);
                puts(STRINGLITERAL);
            }
        }
        seL4_UserContext regs;
        err = seL4_TCB_ReadRegisters(tcb_cap, 0, 0,
                sizeof(seL4_UserContext) / sizeof(seL4_Word), &gregs);
        regs = gregs;
    }
}
```

Listing 10.2: C code for the Microkit `monitor`

Verification stages/steps with artefacts

(These are the same steps as those for the prior example except that the second step is slightly simplified.)

1. Converting into GraphLang $\Longrightarrow$ *control-flow* graph Figure 10.3
2. We note that instead of providing an explicit functional spec, the required

specification in this case of “no undefined behaviour” is being generated by the C parser (see “*C Parser semantics is met*” in Section 4.4.1). Accordingly, this verification step adds the C-Parser-generated spec and then turns the graph into dynamic single-assignment form and acyclic $\Longrightarrow$ expanded control-flow graph Figure 10.4

3. Computing Weakest Precondition (per Section 8.2) $\Longrightarrow$ Logic formula in QF_ABV
4. Running[2] SMT solver `Z3` $\Longrightarrow$ Result = “Yes”

[2]As explained in step 2, for the monitor, there are no function spec invariants that at this stage need to be added manually to the SMT tool as a step 4, unlike in the earlier example of the Arithmetic Sum Section 10.1

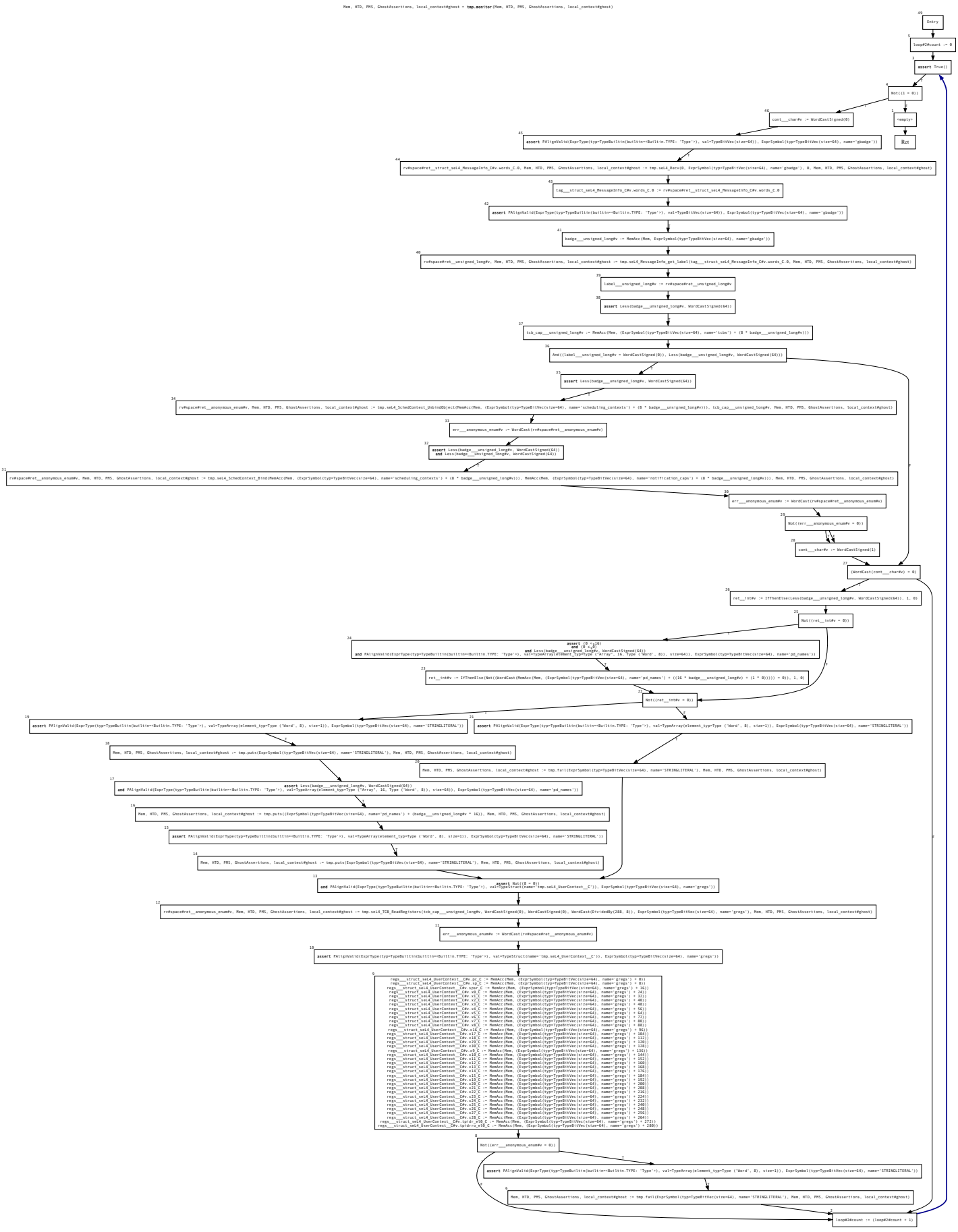


Figure 10.3: Control-flow graph for the Microkit Monitor, as generated from SimplExport

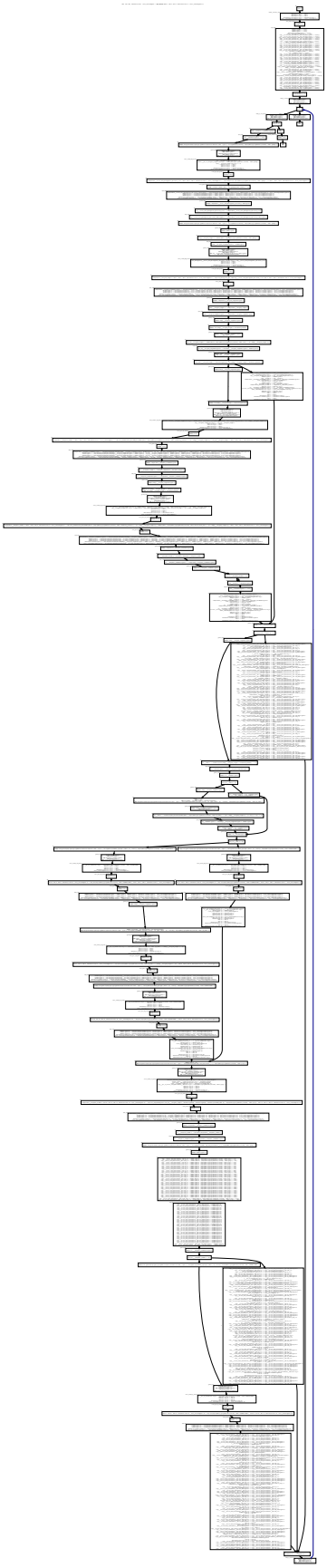


Figure 10.4: Control-flow graph for the Microkit Monitor, with spec and in dynamic single-assignment form and acyclic

Chapter 11

Limitations

11.1 Microkit versions

The verification of the Microkit is based on an implementation version of the Microkit as was available at the time of beginning the project mid 2022 and therefore differs in some key details from the current upstream Microkit version available in the public git repository as of 2024.

11.1.1 Kernel limitations: Non-MCS

The upstream Microkit is built on top of the mixed-criticality scheduling (MCS) seL4 kernel. However, the MCS kernel itself does not yet have a full verification story. In particular, the verified `case-init` loader neither supports MCS nor 64-bit architectures.

In order to obtain a complete verification story for the Microkit, the verification supports not only the upstream Microkit, but also a fork that uses the non-MCS 32-bit kernel. This fork will be redundant once the underlying MCS kernel verification and `case-init` have caught up.

11.1.2 Architecture limitations: 32-bit

Similar caveats hold for the architecture as for the kernel limitations: The verified system initialiser does not presently support 64-bit architectures but only 32-bit architectures, which again, in turn, limits the full verification story of the Microkit.

11.1.3 SDF to CapDL mapping: platform-aware

Ultimately, the functionality for automatically generating CapDL output based on the user-provided system SDF spec should target all platforms supported by the Microkit. However, the Microkit code is by nature platform-aware (as opposed to platform-agnostic), and the Microkit environment currently supports only a few development boards (essentially a limited set of AArch64 boards). The specification may have to be generalised to be platform-agnostic to a greater extent as support for new development boards expands.

A current particular case of such an ’architecture’ limitation relates to the Coupling Invariant not covering memory regions due to specifics of the particular architecture of the seL4 kernel that is used in the Microkit verification.

An other instance is in the context of the Isabelle proof-script in the Microkit proof toolchain (Section 6.1 item (3) and Footnote 2): The proof-script file is well-annotated, and set to succeed automatically on all correct inputs. The translator supports all SDF features, and the verification script can run successfully on all these outputs. However, as the Isabelle session `Types-D` remains incomplete in its port to AArch64 within the seL4 kernel verification, frame caps are excluded from the verification checks. Consequently, the negative aspects of the proof, which demonstrate that the translator does not issue excessive caps beyond the specification's requirements, cannot be implemented at this stage.
This means that, while the verifier proves that no caps were missed by the translator (and hence that no runtime errors or warnings will occur from missing or incorrectly assigned caps), it cannot yet provide security guarantees showing that no unnecessary caps were issued.
Fortuitously, this is a minor and controllable limitation and can be dealt with as the required architecture ports become available. (By way of postscriptum, proofs for the seL4 kernel on AArch64 have been done in 2024, thus opening the path to re-evaluate the closing of this particular gap.)

11.1.4 Build system

As mentioned in Section 9.3, our verified version of the Microkit supports and integrates with the verified CASE initialiser as well as Spinale's new Rust CapDL loader. The latter's key limitation is that it is not a verified initialiser. However it comes with many significant benefits, in particular for build system.

A motivator for using Spinale's new Rust CapDL loader is that it had been designed with the SDK model in mind. The existing CapDL loader written in C required re-compilation every time a system designer would change the capability distribution of their system. In addition, the C CapDL loader was tightly integrated with the rest of the seL4 build system, which meant that trying to fit it in an SDK model posed significant friction. Spinale's new Rust CapDL loader do not exhibit these down-sides.

11.2 Gap in end-to-end proof

11.2.1 Verifying MSpec

The key assurance gap for end-to-end proof is **Proof ③** in Figure 1.4: We still need to demonstrate that MSpec, the kernel specification relevant to the Microkit, is a *provably correct abstraction* of the kernel's abstract specification, ASpec, when we manually derive MSpec from ASpec to make it suitable for SMT solvers.

11.2.2 Haskell to SMT-LIB 2 transcription

In our proof pipeline of the Microkit verification process (Figure 4.2) we note that the transcription of the spec in Haskell to SMT-LIB 2 is a *manual* process. More specifically, it is a 2-step process of first manually transcribing the Haskell spec into Python, and then translating with `Gordian` to SMT-LIB 2.

At the time, the use of Haskell for the formal specification allowed for efficient initial prototyping and development. Meanwhile, the eventual implementation of the Microkit verifier

is SMT-solver based, while Isabelle/HOL is to be used for proofs of properties relating the MSpec with the ASpec. Accordingly, a manual translation of the Haskell specification into Isabelle/HOL is required with, eventually, an *automated* translation from the Isabelle/HOL spec to SMT-LIB 2 that is simple enough to demonstrate the correspondence between these two versions of the specification – as summarised in Figure 11.1.

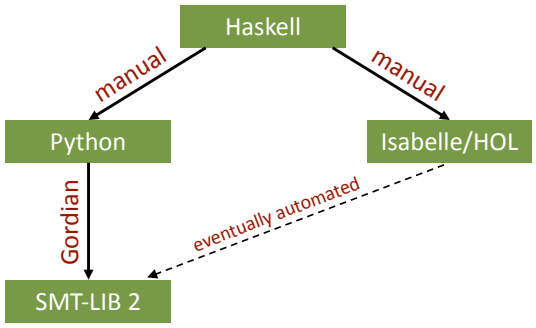


Figure 11.1: Translations of the specifications for the Microkit verification process.

In fact, once the initial specs in Haskell have been translated into SMT-LIB 2 and Isabelle/HOL, as required, these Haskell specs are no longer crucial to the verification process. We will therefore no longer maintain the Haskell specs, and consider the Isabelle/HOL and SMT-LIB 2 specs authoritative, with a simple *automated* translation directly between the two serving to assure their correspondence.

11.3 Threats to validity

There are a number of assumptions on which the proofs reported here are based.

- We prove *functional correctness* only between the induced semantics for the C code and its specification, and so the *compiler and linker need to be trusted*. This gap can be bridged by combining our Microkit tool with the existing seL4 binary-verification toolchain. This toolchain proves that the seL4 binary has the same semantics as the verified C code, and thus ensures that the seL4 kernel proofs apply to the kernel binary. In the same way, that toolchain should be able to extend the verification of `libmicrokit` to its binary code.
- The proof work is done by SMT solvers. We therefore *assume that the SMT solver used is functionally correct* and is invoked correctly. (We have alleviated this concern/gap by having employed a variety of different SMT solvers.)
- Finally, one also has to make *some bottom-level assumptions about the physical world and other code* running in the system. Tackling these have to be left to future work (where possible) or have to be validated by empirical means. If these assumptions are not met, faults can still occur.

  In our case, the assumptions are that the *hardware* works as specified by the manufacturer, the *kernel has been loaded* correctly, and that the *libraries* outside the scope of the verification project, such as `libsel4`, also satisfy the properties stated in their specs. Furthermore, the correct initialisation by `system_initialiser` is assumed unless the verified `case-init` loader is used.

Chapter 12

Achievements and Impact of Verification

12.1 Achievements: What we have proved and delivered

The automated verification of the Microkit implementation proceeds via multiple stages (see Figure 12.1) and reuses a number of tools and libraries that had been developed at the time for the kernel verification effort.

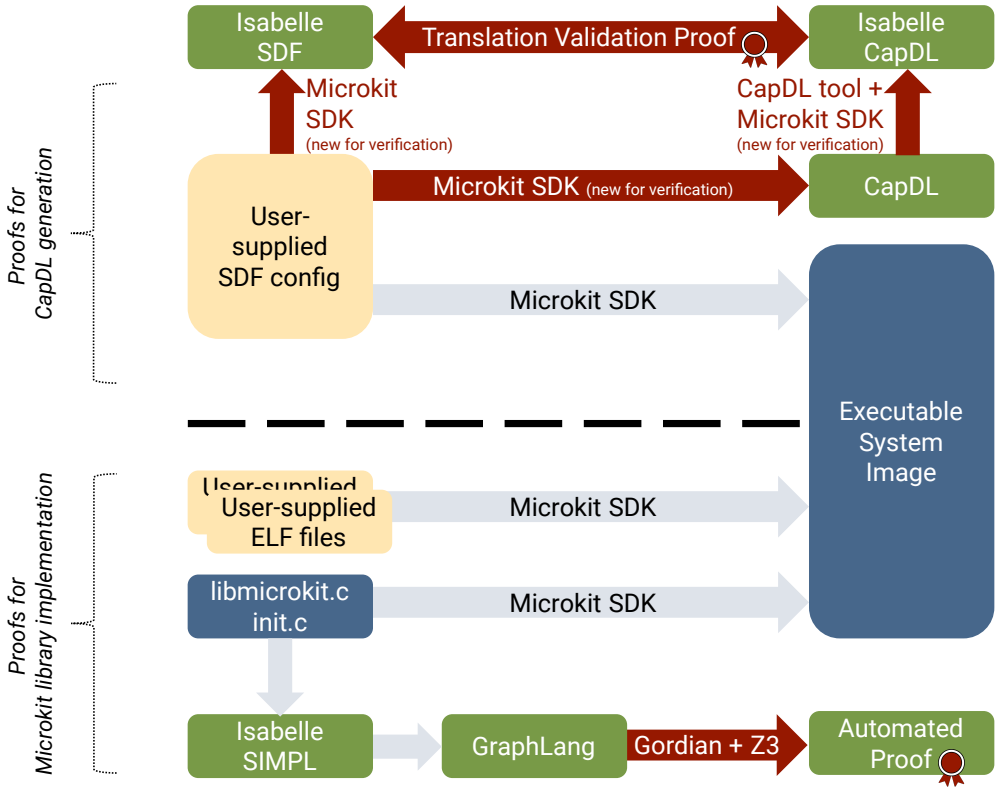


Figure 12.1: Summary of the verification processes for the translation validation of the generated CapDL export and for the automated verification of the Microkit library. In yellow are user-supplied files, in green formal artefacts created as part of the Microkit toolchain and in blue the informal ones. Dark red arrows indicate the new translation and verification tools plus expanded Microkit SDK developed over the course of this verification project, while light grey arrows are the result of prior work. The dashed line separates the two sets of proofs.

Functional correctness

We have verified the Microkit implementation, encompassing both the library (`libmicrokit`) and monitor task (`monitor`) components, using *push-button methods*: Upon successful completion, the `Gordian` verifier confirms that a provided C function is **functionally correct according to its specification**.

In particular, our proof chain with `Gordian` ensures the **absence of Undefined Behaviour**, i.e. of common programming errors, including:

- no null pointer dereferences;
- no incorrect use of dynamic memory during program execution (e.g., no ill-typed or dangling pointers, no out-of-bounds errors);
- no arithmetic overflows and exceptions (e.g., no signed integer overflows, no division-by-zero, no invalid bit shifts, no invalid conversions);
- no other undefined behavior (e.g., not trying to use values of uninitialised local variables).

Furthermore, and importantly, **non-termination is allowed** in the sense that in our specifications for the Microkit components, non-termination is always an allowed behavior. For example, the handler loop is specified to never terminate, and the verifier explicitly confirms this property.

Re-use of `Gordian`

The design and development of our automated verification tool for the task, `Gordian`, has been built with the intention to be re-used in the verification of future projects built within the Microkit framework.

12.2 Impact

Error elimination

The final verification succeeded without finding new bugs or errors in the implementation. However, we identified and eliminated two errors during the specification and verifier development process, improving the overall quality of the code.

As a result, the insights gained during the formal specification and later verification process have already raised the assurance levels of the Microkit and the surrounding ecosystem.

Rust CapDL loader

Moreover, the desire for better assurance led Nick Spinale to develop a new CapDL loader written in Rust [Spinale, 2023]. Since Rust ensures memory safety at compile time using its ownership mechanism and built-in borrow checker, the Rust CapDL loader serves as a safer *drop-in* replacement for the original C CapDL loader – be this for Microkit-based systems or, in fact, for other CapDL-based systems, legacy or otherwise.

Chapter 13

Conclusions

The Microkit verification project has demonstrated that automated verification techniques have potential for reducing the cost of extending seL4's verification into usermode components, as long as they are simple enough. While gaps remain in an end-to-end verification of the Microkit, these can be bridged with established techniques.

The experience from this project make us optimistic about using similar techniques in verifying systems on top of the Microkit, specifically the new Lions OS [Heiser, 2024].

However, the experience has also shown us the need for better tooling. The current `Gordian` approach of inserting verification conditions in the control-flow-graph representation of the program (see Section 8.1) is not sufficiently scalable – the GraphLang representation is more than an order of magnitude larger than the original C program. These annotations should be made in the original program source – not only is the source code a far more compact and readable representation than the control-flow graph, annotating the source will also keep program code and spec together, and thus easier to maintain.

Addressing this issue is subject of on-going research.

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