Prisma: A Tierless Language for Enforcing Contract-Client Protocols in Decentralized Applications (Extended Version)

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Abstract

Decentralized applications (dApps) consist of smart contracts that run on blockchains and clients that model collaborating parties. dApps are used to model financial and legal business functionality. Today, contracts and clients are written as separate programs – in different programming languages – communicating via send and receive operations. This makes distributed program flow awkward to express and reason about, increasing the potential for mismatches in the client-contract interface, which can be exploited by malicious clients, potentially leading to huge financial losses.

In this paper, we present Prisma, a language for tierless decentralized applications, where the contract and its clients are defined in one unit and pairs of send and receive actions that "belong together" are encapsulated into a single direct-style operation, which is executed differently by sending and receiving parties. This enables expressing distributed program flow via standard control flow and renders mismatching communication impossible. We prove formally that our compiler preserves program behavior in presence of an attacker controlling the client code. We systematically compare Prisma with mainstream and advanced programming models for dApps and provide empirical evidence for its expressiveness and performance.

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1 Introduction

dApps enable multiple parties sharing state to jointly execute functionality according to a predefined agreement. The main component is the smart contract – the predefined agreement – that regulates the interaction between the dApp's clients. Client–contract interactions can be logically described by state machines [78, 60, 61, 82] specifying which party is allowed to do what and when.

dApps can operate without centralized trusted intermediaries by relying on a blockchain and its consensus protocol. To this end, a contract is deployed to and executed on the blockchain, which guarantees its correct execution; clients that run outside of the blockchain can interact with the contract via transactions. A key feature of dApps is that they can directly link application logic with transfer of monetary assets. This enables a wide range of correctness/security-sensitive business applications, e.g., for cryptocurrencies, crowdfunding, and public offerings,[1] and make them, at the same time, an attractive target for attackers. The latter is aggrevated by the fact that contracts can be called by any client in the network, including malicious ones that try to force the contract to deviate from the intended behavior [47]. These issues demand programming models that ensure the correctness and security of dApps by-design, especially given the fact that contract code cannot be updated after deployment.

We address correctness and security of the client–contract interaction of dApps by enforcing the execution order of contract code. Deviations from the intended interaction protocols due to implementation errors and/or malicious attacks are indeed a critical threat (besides issues with arithmetic or buffer overflows, etc.). In recent years, there have been several large attacks exploiting flawed program flow control in smart contracts. Most famously, attackers managed to steal around 50 M USD [9, 47] from a decentralized autonomous organization, the DAO. In two attacks on the Parity multi-signature wallet, attackers stole respectively froze cryptocurrencies worth 30 M USD [27] respectively 150 M USD [69].

Programming dApps. The design space of dApp programming languages can be structured along two dimensions – similar to other multi-party applications – based on: (1) how the involved parties (contract and clients) are defined, and (2) how the correctness of their interactions is ensured.

There are two choices for (1): a *local* versus a *global model.* In a *local model*, parties are defined each in a separate *local* program and their interactions are encoded via send and receive effects. In contrast, in a *global model*, parties are defined within one shared program and

[1] 700 K to 2.7 M contracts have been deployed per month between July 2020 and June 2021 [15] on the Ethereum blockchain – the most popular dApps platform [38]. Some dApps manage tremendous amounts of assets, e.g., Uniswap [17] – the largest Ethereum trading platform had a daily trading volume of 0.5 B – 1.5 B USD in June 2021.

interactions are encoded via combined send-and-receive operations with no effects visible to the outside world. The local model is followed by actor systems [18] or by approaches stemming from process calculi [54]; the global model by tierless [71, 37, 31, 32, 49, 79, 80, 72, 83] and choreographic [55, 64, 51] languages.

The local model requires an explicitly specified protocol for governing the inter-process send-receive relations to ensure that every send effect has a corresponding receive effect in an interacting – separately defined – process. In contrast, with a global model, there is no need to separately specify such a protocol. All parties run the same program in lock-step, where a single send-and-receive operation performs a send when executed by one party and a receive on the other party. Due to this encapsulation of the communication effects, there is no non-local information to track – the program's control flow defines the correct interaction and a standard System-F-style type system is sufficient.

Approaches to dApp programming – industrial or research ones – follow a local model, thus rely on explicitly specifying the client–contract interaction protocol. Moreover, the contract and clients are implemented in different languages, hence, developers have to master two technology stacks. The dominating approach in industry uses Solidity [63] for the contract and JavaScript for clients. Solidity relies on developers expressing the protocol using assertions in the contract code, which are checked at run time [13]. Failing to correctly introduce assertions may give parties illegal access to monetary values to the detriment of others [65, 58]. In research, contract languages [39, 33, 74, 75, 35, 34, 66, 25] have been proposed that rely on advanced type systems such as session types, type states, and linear types. There, processes are typed by the protocol (of side-effects such as sending and receiving) that they follow on evaluation, and non-compliant processes are rejected by the type-checker.

The global model has not been explored for dApp programming. This is unfortunate given the potential to get by with a standard typing discipline and to avoid intricacies and potential mismatches of a two-language stack. Our work fills this gap by proposing Prisma – the first language that features a *global programming model* for Ethereum dApps. While we focus on the Ethereum blockchain, we believe our techniques to be applicable to other smart contract platforms.

Prisma. Prisma enables interleaving contract and client logic within the same program and adopts a *direct style (DS)* notation for encoding send-and-receive operations akin to languages with baked-in support for asynchronous interactions, e.g., via async/await [24, 73]. DS addresses several shortcomings with the currently dominant encoding of the protocol's *finite state machines (FSM)* as one method per transition [63, 33, 74, 75, 35, 34].

First, FSM style corresponds to a control-flow graph of basic blocks, which is low-level and more suited to be written by a compiler than by a human – Prisma leaves it to the compiler to map down high-level declarative DS to low-level FSM style. More importantly, ensuring safety in FSM style requires a form of dependent types, as the type of an action depends on the current state. Second, with FSM style, the contract is a passive entity whose execution is driven by clients. The DS encoding, in contrast, allows the contract to actively ask clients for input. The latter style naturally fits an execution model where a dominant acting role controls the execution and diverts control to other parties when their input is needed, which matches well the dApp setting.

Overall, Prisma relieves the developer from the responsibility of correctly managing distributed, asynchronous program flows and the heterogeneous technology stack. Instead, the burden is put on the compiler, which distributes the program flow by means of selective continuation-passing-style (CPS) translation and defunctionalisation and inserts guards against malicious client interactions. We needed to develop a CPS translation for the code that runs on the Ethereum Virtual Machine (EVM) since the EVM has no built-in support for concurrency primitives to suspend execution and resume later, which could be used to implement asynchronous communication. While CPS translations are well-known, we cannot use them out-of-the-box because the control flow is interwoven with distribution in our case. Given that CPS translations reify control flows into values, a CPS translation that does not take distribution into account would allow malicious clients to force the contract to deviate from the intended control flow by passing a spoofed value to the contract. Thus, it is imperative to prove correctness of our *distributed CPS translation* and ensure control-flow integrity of the contract. We formally define the compilation steps and prove them sound and, by doing so, eliminate the risk of programmers implementing unsafe interactions that can potentially be exploited.

Contributions. We make the following contributions:

1. We introduce Prisma[2], a global language for tierless dApps with direct-style client–contract interactions and explicit access control, implemented as an embedded DSL in Scala. Further, Prisma automatically enforces the correct program flow (Section 2).
2. A core calculus, MiniPrisma, which formalizes both Prisma and its compiler, as well as a proof that our compiler preserves the source program behavior in presence of an attacker that controls the client code (Section 3).
3. Case studies which show that Prisma can be used to implement common applications without prohibitive performance overhead (Section 5).
4. A comparison of Prisma with a session type and a type state smart contract programming language and the mainstream programming Solidity/JavaScript model. (Section 6).

2 Prisma in a Nutshell

We present Prisma by the example of a TicTacToe game, demonstrating that client and contract are written in a single language, where protocols are expressed by control flow (instead of relying on advanced typing disciplines) and enforced by the compiler.

Example. TicTacToe is a two-player game over a 3×3 board in which players take turns in writing their sign into one of the free fields until all fields are occupied, or one player wins by owning three fields in any row, column, or diagonal. The main transaction of a TicTacToe dApp is `move(x,y)` used by a player to occupy field `(x,y)`. A `move(x,y)` is valid if it is the sender's turn and `(x,y)` is empty. Before the game, players deposit their stakes, and after the game, the stakes are paid to the winner.

Fig. 1 depicts possible control flows with transitions labeled by client actions that trigger them. Black arrows depict intended control flows. The dApp starts in the funding state where both parties deposit stakes via $\mathit{Fund}(c)$. Next, parties execute $\mathit{Move}(x, y)$ until one party wins or the game ends in a draw. Finally, any party can invoke a payout of the stakes via $\mathit{Payout}()$.[3] Red dashed arrows illustrate the effects of a mismanaged control flow: a

[2] Prisma implementation and case studies are publicly available: `https://github.com/stg-tud/prisma` (anonymized for double-blind review)
[3] We omit handling timeouts on funding and execution for brevity.

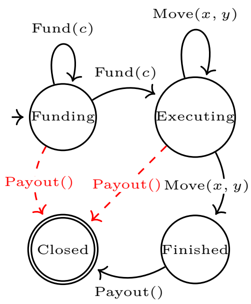


Figure 1 TicTacToe control flow.

Table 2 Location annotations.

Annotations	Description
`@co`	on contract
`@cl`	on clients
`@co @cl`	independent copies on clients and contract
`@co @cross`	on contract, but also accessible by client
`@cl @cross`	(illegal combination)

```
@prisma object TicTacToeModule {

  @co @cl case class UU(x: U8, y: U8)

  class TicTacToe(
    val players: Arr[Address],
    val fundingGoal: Uint) {

    // u8 is an unsigned 8-bit integer
    @co @cross var moves: U8 = "0".u8
    @co @cross var winner: U8 = "0".u8
    @co @cross val board: Arr[Arr[U8]] =
      Arr.ofDim("3".u, "3".u)

    @co def move(x: U8, y: U8): Unit =
      { /* ... */ }
    @cl def updateBoard(): Unit =
      { /* ... */ }
    @cl def fund(): (U256, Unit) =
      (readLine("How much?").u, ())
    @cl def decideMove(): (U256, UU) =
      ("0".u, UU(readLine("x-pos?"),
                 readLine("y-pos?"))
    @co val init: Unit = {
      while (balance() < FUNDING_GOAL) {
        awaitCl(_ => true) { fund() }
      }
      while (moves < "9".u && winner == "0".u) {
        val pair: UU = awaitCl(a =>
          a == players(moves % "2".u)) { decideMove() }
        move(pair.x, pair.y)
      }
      awaitCl(a => true) {
        readLine("Press (enter) for payout")
        ("0".u, ())
      }
      if (winner != "0".u) {
        players(winner - "1".u).transfer(balance())
      }
      else {
        players("0".u).transfer(balance() / "2".u)
        players("1".u).transfer(balance()) // remainder
      }
    }
  }
}
```

Figure 3 TicTacToe dApp.

malicious player could trigger a premature payout preventing the counterpart to get financial gains.

Tierless dApps. Prisma is implemented as a DSL embedded into Scala. Hence, the Prisma code we show resembles Scala. Prisma interleaves contract and client logic within the same program. Annotations `@co` and `@cl` explicitly place declarations on the contract and on the client, resp. (cf. Tab. 2). A declaration marked as both `@co` and `@cl` has two copies. For security, code placed in one location cannot access definitions from the other — an attempt to do so yields a compile-time error. Developers can overrule this constraint to enable clients to read contract variables or call contract functions by combining `@co` with `@cross`. Combining `@cl` with `@cross` is not allowed – information can only flow from client to contract as part of a client–contract interaction protocol.

There are three kinds of classes. *Located classes* are placed in one location (annotated with either `@co` or `@cl`); they cannot contain located members (annotated with either `@co` or `@cl`) and cannot cross the client–contract boundary, e.g., cannot be passed to or returned from `@cross` functions. *Portable classes* are annotated with both `@co` and `@cl`. Their instances can be passed to and returned from `@cross` functions; they must not contain mutable fields. *Split classes* have no location annotation; their instances live partly in both locations; they cannot be passed to or returned from `@cross` functions and their members must be located.

Prisma code is grouped into modules. While client declarations can use and be used from standard (non-Prisma) Scala code, contract declarations are not accessible from Scala code, and can only reference contract code from other Prisma modules (because contract/client code lives in different VMs).

For illustration, consider the TicTacToe dApp (Fig. 3). The `TicTacToeModule` (Line 1) – modules are called `object` in Scala – contains a portable class `UU` (Line 3) and a split class `TicTacToe` (Line 5). Variables `moves`, `winner`, `board` (Lines 10, 11, 13) are placed on the contract and can be read by clients (`@co @cross`). The `updateBoard` function (Line 17) is placed on the client and updates client state (e.g., client's UI). The `move` function (Line 15) is placed on the contract and changes the game state (`move`). `move` is not annotated with `@cross`, because `cross` is intended for functions that do not change contract state and can be executed out-of-order without tampering with the client–contract interaction protocol.

Scala only has literals signed integers and signed longs, which are uncommon for use in Ethereum. Prisma provides portable unsigned and signed integers for power-of-two bitsizes between 2^3 to 2^8, with common arithmetic operations, e.g., `"0".u8` is an unsigned 8-bit integer of value 0 (Line 10).

Encoding client–contract protocols. In Prisma, a client-contract protocol is encoded as a split class containing dedicated `awaitCl` expressions for actively requesting and awaiting messages from specific clients and standard control-flow constructs. Hence, creating a new contract instance corresponds to creating a new instance of a protocol; once created, the contract instance actively triggers interactions with clients.

The `awaitCl` expressions have the following syntax:

```
def awaitCl[T](who: Addr => Bool)(body: => (Ether, T)): T
```

They take two arguments. The first (`who`) is a predicate used by clients to decide whether it is their turn and by the contract to decide whether to accept a message from a client. This is unlike Solidity, where a function may be called by any party by default. By forcing developers to explicitly define access control, Prisma reduces the risk of human failure. The second argument (`body`) is the expression to be executed by the client. The client returns a pair of values: the amount of Ether and the message to the contract. The former can be accessed by the contract via the built-in `value`, the latter is returned by `awaitCl`. Besides receiving funds via `awaitCl`, a contract can also check its current balance (`balance()`), and transfer funds to an account (`account.transfer(amount))`).

For illustration, consider the definition of `init` on the right-hand side of Fig. 3, Line 24. It defines the protocol of `TicTacToe` as follows. From the beginning of `init` the flow reaches `awaitCl` in Line 26 where the contract waits for clients to provide funding (by calling `fund`). Next, the contract continues until `awaitCl` in Line 29 and clients execute `decideMove` (Line 21) until the game ends with a winner (`winner != 0`) or a draw (`moves >= 9`). At this point, – `awaitCl` in Line 33 – any party can request a payout and the contract executes to the end. The example illustrates how direct-style `awaitCl` expressions and the multitier model enable encoding multiparty protocols as standard control flow, with protocol phases corresponding to basic blocks between `awaitCl` expressions.

Compiling Prisma to Solidity. Prisma's compiler splits contract methods that contain `awaitCl` expressions into one method per `awaitCl`, each taking the result from the last `awaitCl` as

```
1 contract TTT { State phase = Funding; /*...*/
2   function Fund() public {
3     require(phase==Funding);
4     /*...*/; if (!(balance < FUNDING_GOAL)) phase=Exec }
5   function Move(int x, int y) public {
6     require(phase==Exec && sender==players(moves % 2));
7     /*...*/; if (!(moves < "9".u && winner == "0".u)) phase=Finished }
8   function Payout() public {
9     require(phase==Finished); /*...*/; phase=Closed }
```

Figure 4 Generated Solidity

an argument. The compiler also hardens the program flow against malicious clients. Once deployed, a contract is public and can be messaged by clients not generated by Prisma – hence, we cannot assume that clients will actually execute the body passed to them by an `awaitCl` expression. To cope with malicious clients trying to tamper with the control flow of the contract, the compiler generates code to store the current phase before giving control to the client and to reject methods invoked by wrong clients or in the wrong phase.

For illustration, the code generated from `init` in Fig. 3 is schematically shown in Fig. 4. Three methods are generated – one for each `awaitCl` expression in `init` – corresponding to phases in the logical protocol (Fig. 1). The code up until the first `awaitCl` (Line 26) is placed in the constructor of the generated contract, which ends by setting the active phase to `Funding`. The code between the first and the second `awaitCl` either loops back to the first `awaitCl` or continues to the second one (Line 29), it is placed in the `Fund` method that requires the phase to be `Funding`, and may change it to `Exec`, if the loop condition fails. Similarly, the method `Move` is generated to contain the loop between the second and the third `awaitCl` (Line 33); and the method `Payout` contains the code from the third `awaitCl` to the end of `init`. Only the second `awaitCl` contains a (non-trivial) access control predicate, which shows up as an additional assertion in the body of `Move` (Line 6).

3 Compilation and its Correctness

Prisma is compiled to Solidity and Scala. An important part of our compiler is CPS translation. Using a standard CPS translation alone, however, is not sufficient because the control flow is intertwined with distribution. We need to consider that, while malicious clients cannot change the code executed by the contract, they can run different code on the client side than what was assigned to them. Therefore, clients can pass unintended continuations to the contract. To ensure the correctness of our compiler, we formalize a selective CPS translation, defunctionalisation toghether with the process of guarding against unintendend jumps. Assuming the worst case, i.e., all clients are malicious and replace their code completely executing arbitrary code instead, we prove guarding contract code is sufficent to guarantee that behavior of the contract execution is preserved by compilation. The compilation process is key in hiding the complexity of enforcing distributed program flow from the developer – hence, correctness is critical.

- We present the core calculus (Section 3.1) MiniPrisma$_*$ – a hybrid language that includes elements of both the source (MiniPrisma$_s$) and the compilation target (MiniPrisma$_t$), while abstracting over details of both Scala and Solidity. We define a hybrid language because the source and the target share many constructs – the hybrid language allows us to focus on how the differences are compiled.
- We model the compiler as a sequence of steps that transform MiniPrisma$_*$ programs via several intermediate representations.

- We define MiniPrisma$_*$ semantics as a reduction relation over configurations consisting of traces of evaluation events and expressions being evaluated. We distinguish between a good semantics, which evaluates the program in the usual way, and a bad semantics, which models attackers by ignoring client instructions and producing arbitrary values that are sent to the contract.
- We prove secure compilation by showing that the observable behavior of the programs before and after compilation is equivalent. We capture the observable program behavior by the trace of events generated during program evaluation (as guided by the semantic definition) and show trace equivalence of programs before and after compilation.

In more concrete terms, the last point guarantees the integrity for the control flow of the contract in the presence of malicious clients. Note that compiler correctness in a traditional sense (i.e., preservation of the semantics through compilation) is not sufficient in the presence of malicious attackers that can tamper with parts of the code.

More specifically, the compiler preserves control flow by inserting additional guards in places where the basic blocks of the program have been separated and distributed onto different computers by the CPS translation. This ensures that the control flow among the parties is guaranteed not to change after compilation. In particular, we consider that malicious attackers may tamper with certain parts of the code (the client code). Therefore the guards have to be placed on the contract side. Our guarantees hold despite clients tampering with the code.

For example, in the source code of the TicTacToe game (Section 2), it is immediately visible that the players take alternating turns. Yet, the compiled code is separated and distributed into small chunks of code. Parties execute a chunk and then wait for other parties, which will decide on a move that influences where, when, and how to proceed execution. Our proof shows that we still ensure that control flow of the contract of the original program even under attack.

3.1 Syntax

The syntax of MiniPrisma$_*$ (Fig. 5), has three kinds of identifiers id, i, j, from unspecified sets of distinct names. Pure identifiers id are for function arguments and let bindings; mutable variables i are for heap variable assignment and access. In the target program, mutable variables j ($\mathsf{who}, \mathsf{state}, \mathsf{clfn}, \mathsf{cofn}$) generated by the compiler can also appear. We call compiler-generated identifiers *synthetic*. Normal identifiers are separated from synthetic ones to distinguish compiler generated and developer code. Definitions d and definitions for synthetic identifiers b are semicolon-separated lists of declarations that assign values to variables and annotate either the contract or the client location. Each program P consists of definitions d and synthetic definitions b followed by the main contract expression m. Program P corresponds to a single Prisma split class, d and b to methods and generated methods, and m to a constructor containing the initialisation of its class members (such as the body of `init`, Fig. 3).

Constants c are unsigned 256 bit integer literals and built-in operators. MiniPrisma$_*$ supports tuples introduced by nesting pairs (::) and eliminated by pattern matching. Tuples allow multiple values to cross tiers in a single message. Values v are constants, value pairs, and lambdas. Patterns x are constants, pattern pairs, and variables. Expressions e are constants, expression pairs, lambdas, variables, variable accesses/assignments, bindings and function applications.

$$id \in ID \qquad i \in I \qquad j \in \{\mathsf{who}, \mathsf{state}, \mathsf{clfn}, \mathsf{cofn}\}$$

$$\begin{array}{llll}
\text{(definition)} & d & ::= & \text{@co this}.i = v;\ d \\
 & & \mid & \text{@cl this}.i = v;\ d \mid () \\
\text{(synthetic definition)} & b & ::= & \text{@co this}.j = v;\ b \\
 & & \mid & \text{@cl this}.j = v;\ b \mid () \\
\text{(program)} & P & ::= & d;\ b;\ m \\
\text{(constant)} & c & ::= & 0 \mid 1 \mid 2 \mid \ldots \mid \mathsf{true} \mid \mathsf{false} \\
 & & \mid & () \mid \&\& \mid + \mid == \mid < \mid \mathsf{try} \\
 & & \mid & \gg\!= \mid \mathsf{trmp} \mid \mathsf{Done} \mid \mathsf{More} \\
\text{(value)} & v & ::= & c \mid v :: v \mid x \to e \\
\text{(pattern)} & x & ::= & c \mid x :: x \mid id \\
\text{(expression)} & e & ::= & c \mid e :: e \mid x \to e \mid id \\
 & & \mid & \mathsf{this}.i \mid \mathsf{this}.i := e \\
 & & \mid & \mathsf{this}.j \mid \mathsf{this}.j := e \\
 & & \mid & x{=}e;\ e \mid e\ e \\
\text{(main expression)} & m & ::= & c \mid m :: m \mid x \to e \mid id \\
 & & \mid & \mathsf{this}.i \mid \mathsf{this}.i := m \\
 & & \mid & \mathsf{this}.j \mid \mathsf{this}.j := m \\
 & & \mid & x{=}m;\ m \mid m\ m \\
 & & \mid & \mathsf{awaitCl}_s(e, () \to e) \mid \mathsf{awaitCl}_t(c, () \to e)
\end{array}$$

Figure 5 $\mathsf{MiniPrisma}_*$ syntax.

$$\begin{array}{rcl}
m_0\ c\ m_1 & = & c(m_0, m_1) \\
(m_0, \ldots, m_n) & = & m_0 :: \ldots :: m_n :: () \\
m_0;\ m_1 & = & () = m_0;\ m_1 \\
\mathsf{assert}(m_0);\ m_1 & = & \mathsf{true} = m_0;\ m_1 \\
x \leftarrow e_1;\ m_2 & = & x = \mathsf{awaitCl}_t(() \to e_1);\ m_2 \\
\mathsf{if\ let}\ x = m_1\ \mathsf{then}\ e_2\ \mathsf{else}\ e_3 & = & \mathsf{try}(m_1, x \to e_2, () \to e_3)
\end{array}$$

Figure 6 Syntactic sugar.

Main expressions m may further contain remote client expressions, embedding client code into contract code and waiting for its result. The source client expression $\mathsf{awaitCl}_s(e,\ () \to e)$ can be answered by any client whose address fulfills the predicate specified as first argument. $\mathsf{awaitCl}_s$ corresponds to direct-style remote access via `awaitCl` in Prisma. We use the syntax form $\mathsf{awaitCl}_t(c,\ () \to e)$ to model the execution of code e on the specified client c. $\mathsf{awaitCl}_t$ has no correspondence in the source syntax. Our compilation first splits the predicate from the source client expressions into a separate access control guard. Then, it eliminates client expressions, turning the contract into a passive entity that stops and waits for client input.

We now map the hybrid language $\mathsf{MiniPrisma}_*$ to the source and target languages, $\mathsf{MiniPrisma}_s$ and $\mathsf{MiniPrisma}_t$. $\mathsf{MiniPrisma}_s$ has all expressions of $\mathsf{MiniPrisma}_*$, except those that contain $\gg\!=$ (bind), trmp (trampoline), Done, More, $\mathsf{awaitCl}_t$, or synthetic identifiers j. $\mathsf{MiniPrisma}_t$ has all expressions of $\mathsf{MiniPrisma}_*$ except those that contain $\mathsf{awaitCl}_s$, $\mathsf{awaitCl}_t$, $\gg\!=$. $\gg\!=$ and $\mathsf{awaitCl}_t$ may neither appear in source nor target programs; the former is only an intermediate compiler construct, the latter is used to track of the current evaluating location.

$$
\begin{array}{llll}
\text{(frame)} & F & ::= & \mathsf{awaitCl}_s(\square, () \to e) \mid \square\ e \mid e\ \square \\
 & & \mid & \square :: e \mid e :: \square \mid x = \square;\ e \mid x = e;\ \square \\
 & & \mid & \mathsf{this}.i := \square \mid \mathsf{this}.j := \square \\
\text{(atom)} & a & ::= & \mathsf{this}.i \mid \mathsf{this}.j \mid c \mid id \mid x \to e \\
\text{(context)} & E & ::= & \square \mid E :: m \mid v :: E \mid E\ m \mid v\ E \\
 & & \mid & x = E;\ m \mid \mathsf{this}.i := E \mid \mathsf{this}.j := E \\
\text{(event)} & p & ::= & \mathsf{wr}(c, i, v)\ p \mid \mathsf{msg}(c, v)\ p \mid () \\
\text{(synthetic events)} & q & ::= & \mathsf{wr}(c, j, v)\ q \mid () \\
\text{(configuration)} & C & ::= & p;\ q;\ c;\ m
\end{array}
$$

Figure 7 Frames, Events and configurations.

$$
\begin{array}{lcl}
init_A(d; b; m) & = & init_A(d; b);\ 0;\ m \\
init_A(d; b) & = & (wr(0, i, v) \mid \forall\ (@\mathsf{co}\,\mathsf{this}.i = v) \in d) \\
 & & (wr(0, j, v) \mid \forall\ (@\mathsf{co}\,\mathsf{this}.j = v) \in b) \\
 & & (wr(c, i, v) \mid \forall\ (@\mathsf{cl}\,\mathsf{this}.i = v) \in d, \quad c \in A) \\
 & & (wr(c, j, v) \mid \forall\ (@\mathsf{cl}\,\mathsf{this}.j = v) \in b, \quad c \in A)
\end{array}
$$

Figure 8 Initialization.

Syntactic sugar. In Fig. 6, we define some syntactic sugar to improve readability. We use infix binary operators and tuple syntax for nested pairs ending in the unit value (); we elide the let expression head for let bindings matching (), $\mathsf{assert}(x)$ is a let binding matching true; we use monadic syntax for let bindings of effectful expressions; $\mathsf{if\ let}\ x = m\ \mathsf{then}\ e\ \mathsf{else}\ e$ is the application of the built-in try function.

Events and configurations. In Fig. 7, we define left-to-right evaluation contexts E [45]; and compilation frames F [70], such that every expression decomposes into a frame-redex pair $F\ e$ or is an atom a. Events p and q are lists that capture the observable side-effects of evaluating expressions. They are either (a) state changes $\mathsf{wr}(c, i, v)$ and $\mathsf{wr}(c, j, v)$, from the initial definitions or variable assignment, where i and j are the variable being assigned, c the location, and v the assigned value, or (b) client-to-contract communication $\mathsf{msg}(c, v)$, where c is the address of the client and v the sent value. Configurations $C = p; q; cm$, represent a particular execution state, where p (and q) are traces of normal (and synthetic) events produced by the evaluation, c is the evaluating location, and m is the expression under evaluation.

Initialization. Initialization in Fig. 8 generates the initial program configuration, which models the decentralized application with a single contract and multiple clients. We model a fixed set of clients A interacting with a contract. The initialization of a program $d; b; m$ to a configuration $p; q; 0; m$ leaves the expression m untouched and generates a list of events – one write event for each normal and synthetic definition. Location 0 represents the contract.

3.2 Compilation

The compiler eliminates language features not supported by the compilation target one by one, lowering the abstraction level from (1) *direct style communication (DS)* – which needs language support for !-notation [5] – through the intermediate representations of (2) *monadic*

$$
\begin{aligned}
&mnf'(d;b;m) && = d;b;\mathsf{trmp}(mnf((\text{Done},m)))\\
&mnf(F\ e) && = assoc(id_0{=}e;\ mnf(F\ id_0))\\
&mnf(a) && = a\\
&assoc(x_0{=}(x_1{=}m_1;\ m_0);\ m_2) && = assoc(x_1{=}m_1;\ assoc(x_0{=}m_0;\ m_2))\\
&assoc(m) && = m
\end{aligned}
$$

Figure 9 Monadic normal form transformation.

normal form (MNF) – which needs support for do-notation [2] – and (3) *continuation-passing style (CPS)* – which needs higher-order functions – to (4) explicitly encoding *finite state machines (FSM)* – for which first-order functions suffice. In the following, we provide an intuition for the compiler steps and subsequently their formal definitions.

First, the compilation steps *mnf* and *assoc* transform DS remote communication $\mathsf{awaitCl}e$ to variable bindings ($x \leftarrow e$) and nested let bindings are flattened such that a program is prefixed by a sequence of let expressions. Second, step *guard* generates access control guards around client expressions to enforce correct execution even when clients behave maliciously. Third, step *cps* transforms previously generated let bindings for remote communication ($x \leftarrow e_1;\ m_2$) to monadic bindings $e_1 \ggg x \to m_2$. Fourth, step *defun* transforms functions into data structures that can be sent over the network and are interpreted by a function (i.e., an FSM) on the other side. Compared to standard defunctionalization, we handle two more issues. First, we defunctionalize the built-in higher-order operator ($\ggg$) by wrapping the program expression into a call to a trampoline $\mathsf{trmp}(...)$ and transforming the bind operator $(... \ggg x \to ...)$ to the $(\mathsf{More}, ..., ...)$ data structure; the trampoline repeatedly interprets the argument of More until it returns Done instead of More signaling the program's result. Second, we keep contract and client functions separate by generating separate synthesized interpreter functions, called cofn and clfn, thereby splitting the code into the parts specific to contract and client.

MNF transformation (Fig. 9). The mnf' function wraps the main expression m into a call to the trampoline with the pair (Done, m) – signaling the final result – as argument. Then, *mnf* transforms expressions recursively, binding sub-expressions to variables, resulting in a program prefixed by a sequence of let bindings. As recursive calls to *mnf* may return chains of let bindings, we apply *assoc* to produce a flat chain of let bindings. Given a let binding, whose sub-expressions are in MNF, associativity recursively flattens the expression, by moving nested let bindings to the front, $(...\ (...\ m_0;\ m_1);\ m_2 = ...\ m_0;\ (...\ m_0;\ m_2))$, creating a single MNF expression (i.e., *assoc* is composition for MNF terms).

Guarding (Fig. 10). We insert access-control guards for remote communication expressions $\leftarrow_s$ to enforce (i) the execution order of contract code after running the client expression and (ii) that the correct client invokes the contract continuation. The transformation sets the synthetic variable state to a unique value before the client expression, and stores the predicate to designate valid clients in the synthetic variable who. After the client expression, the generated code asserts that the contract is in the same state, and checks that the sender fulfills the predicate. The assertion trivially holds in the sequential execution of the source language, but after more compilation steps the client will be responsible for calling the correct continuation on the contract. Since client code is untrusted, the contract needs to ensure that only the correct client can invoke only the correct continuation.

$$
\begin{array}{lcl}
\mathit{guard}'(d; b; \mathsf{trmp}(m)) & = & d; b; \mathsf{trmp}(\mathit{guard}(m)) \\
\mathit{guard}\begin{pmatrix} x \leftarrow_s (e_0, () \rightarrow e_1); \\ m_2 \end{pmatrix} & = & \begin{pmatrix} \mathsf{this.who} := e_0;\ \mathsf{this.state} := c; \\ x \leftarrow_s (() \rightarrow \mathsf{true}, () \rightarrow e_1); \\ \mathsf{assert}(\mathsf{this.state} == c\ \&\& \\ \qquad \mathsf{this.who}(\mathsf{this.sender})); \\ \mathsf{this.state} := 0;\ \mathit{guard}(m_2) \end{pmatrix} \\
 & & \text{where } c \text{ fresh} \\
\mathit{guard}(x = e_0;\ m_1) & = & x = e_0;\ \mathit{guard}(m_1) \\
\mathit{guard}(m) & = & m
\end{array}
$$

Figure 10 Guarding.

$$
\begin{array}{lcl}
\mathit{cps}'(d; b; \mathsf{trmp}(m)) & = & d; b; \mathsf{trmp}(\mathit{cps}(m)) \\
\mathit{cps}(x \leftarrow_s (() \rightarrow \mathsf{true}, e_0);\ m_1) & = & e_0 \gg\!= (x \rightarrow \mathit{cps}(m_1)) \\
\mathit{cps}(x = e_0;\ m_1) & = & x = e_0;\ \mathit{cps}(m_1) \\
\mathit{cps}(m) & = & m
\end{array}
$$

Figure 11 Continuation-passing style transformation.

$$
\begin{array}{lcl}
\mathit{defun}'(d; b; e) & = & \mathit{defun}(d; \mathit{coclfn}(b, \mathit{id}, \mathsf{assert}(\mathsf{false}), \mathsf{assert}(\mathsf{false})); e) \\
 & & \text{where} \quad \mathit{id}\ \text{fresh} \\
\mathit{defun}\begin{pmatrix} d; \mathit{coclfn}(b, \mathit{id}, \\ \quad e_{1,alt}, \\ \\ \quad e_{2,alt} \\ \quad); \\ (() \rightarrow e_1) \gg\!= (x \rightarrow e_2) \end{pmatrix} & = & \begin{pmatrix} d; \mathit{coclfn}(b, \mathit{id}, \\ \quad \mathsf{if\ let}\ (c :: \mathit{fv}(() \rightarrow e_1)) = \mathit{id} \\ \qquad \mathsf{then}\ e_1\ \mathsf{else}\ e'_{1,alt}, \\ \quad \mathsf{if\ let}\ (c :: x :: \mathit{fv}(x \rightarrow e'_2)) = \mathit{id} \\ \qquad \mathsf{then}\ e'_2\ \mathsf{else}\ e'_{2,alt}); \\ (\mathsf{More},\ c :: \mathit{fv}(() \rightarrow e_1),\ c :: \mathit{fv}(x \rightarrow e'_2)) \end{pmatrix} \\
 & & \text{where} \quad c\ \text{fresh} \\
 & & \text{and} \quad d; \mathit{coclfn}(b, \mathit{id}, e'_{1,alt}, e'_{2,alt}); e'_2 = \\
 & & \qquad \mathit{defun}(d; \mathit{coclfn}(b, \mathit{id}, e_{1,alt}, e_{2,alt}); e_2) \\
\mathit{defun}\begin{pmatrix} d; \mathit{coclfn}(b, \mathit{id}, \\ \quad e_{1,alt}, e_{2,alt}); \\ x = e_0;\ e_1 \end{pmatrix} & = & d; \mathit{coclfn}(b, \mathit{id}, e_{1,alt}, e_{2,alt}); x = e_0;\ \mathit{defun}(e_1) \\
\mathit{defun}\begin{pmatrix} d; \mathit{coclfn}(b, \mathit{id}, \\ \quad e_{1,alt}, e_{2,alt}); \\ e \end{pmatrix} & = & d; \mathit{coclfn}(b, \mathit{id}, e_{1,alt}, e_{2,alt}); e \\
\mathit{coclfn}(b, \mathit{id}, e_{1,alt}, e_{2,alt}) & = & \mathsf{@cl\ this.clfn} = \mathit{id} \rightarrow e_{1,alt}; \\
 & & \mathsf{@co\ this.cofn} = \mathit{id} \rightarrow e_{2,alt}; b
\end{array}
$$

Figure 12 Defunctionalization.

CPS transformation (Fig. 11). The *cps* transformation turns the chains of let bindings produced by *mnf* into CPS. The chain contains three cases of syntax forms: (1) monadic binding ($x \leftarrow ...;\ m_1$), (2) let binding ($x = e_0;\ m_1$), or (3) final expression. For (1), *cps* replaces the monadic binding with an explicit call to the bind operator ($... \gg= (x \to cps(m_1))$). For (2) and (3), *cps* recurses into the tail of the chain. This resembles do-notation desugaring (e.g., in Haskell).

Defunctionalization (Fig. 12). The *defun* function transforms the chains of let bindings and bind operators produced by *cps*, which contains three cases of syntax forms: (1) a bind operator ($e_1 \gg= e_2$), or (2) a let binding ($x = e_1;\ e_2$), or (3) the final expression. For (1), e_1 and e_2 are replaced by data structures that contain values for the free variables in e_1 and e_2 and are tagged with a fresh ID. The body of the expression is lifted to top-level synthetic definitions. For this, *defun* modifies the synthetic definitions b by extracting the body $e_{1,alt}$ of the synthetic clfn definition and the body $e_{2,alt}$ of cofn, and by adding an additional conditional clause to these definitions. The added clause answers to requests for a given ID with evaluating the original expression. For (2) and (3), *defun* recurses into the expressions.

After defunctionalization, lambdas $x \to e_0$ are lifted and assigned a top-level identifier id_0 and lambda applications, $id_0(e_1)$, are replaced with calls to a synthesized interpreter function $\mathsf{fn}(id_0, e_1)$. The latter branches on the identifier and executes the code that was lifted out of the original function.

Compiling. The *comp* function composes the compiler steps after *mnf*:

$$
\begin{aligned}
comp &= defun \circ cps \circ guard \\
comp' &= defun' \circ cps' \circ guard'
\end{aligned}
$$

3.3 Semantics

We model the semantics as a reduction relation over configurations $p; q; c; m \to p'; q'; c'; m'$. Location $c = 0$ denotes contract execution, otherwise execution of client of address c. We distinguish good ($\to_g$) and bad ($\to_b$) evaluations (Fig. 14 and 15); shared rules are in black, without subscript ($\to$).

Attacker model. Attackers can control an arbitrary number of clients and make them send arbitrary messages. Hence, the bad semantics can answer a request to a client with an arbitrary message from an arbitrary *id*. We use evaluation with bad semantics to show that our compiler enforces access control against malicious clients.

Good evaluations of client expressions in the source language (Rgs) and calls to the trampoline in the target language (Rgt) reduce to a client that fulfils the given predicate. We require that predicates evaluate purely. Hence, p and q do not change in the evaluation. On the other hand, bad evaluation of client expressions in the source language (Rbs) and of calls to the trampoline in the target (Rbt) ignore the predicate, choosing an arbitrary client. The trampoline ends, when it reaches Done (Rt). Further, after chosing a client to evaluate, the good evaluation (Rg) continues to reduce the client expression to a value, while the bad evaluation (Rb) replaces the expression e with a (manipulated) arbitrary value v'. Both evaluations (Rg, Rb) emit the message event $\mathsf{msg}(c, v)$ and an assignment to the special variable sender, when a client expressions is reduced to a value v, to record the client–contract interaction.

$$
\begin{aligned}
[c \mapsto c] &= [] \\
[id \mapsto v] &= [id \mapsto v] \\
[(e_0 :: e_1) \mapsto (e_0' :: e_1')] &= [e_0 \mapsto e_0'] \cdot [e_1 \mapsto e_1']
\end{aligned}
$$

Figure 13 Pattern matching.

(RGS)
$$\frac{p; q; 0; v(c) \;\to^* p; q; 0; \mathsf{true}}{p; q; 0;\ \mathsf{awaitCl}_s(v, () \to e) \;\to_g p; q; 0;\ \mathsf{awaitCl}_t(c, () \to e)}$$

(RBS)
$$p; q; 0;\ \mathsf{awaitCl}_s(v, () \to e) \;\to_b p; q; 0;\ \mathsf{awaitCl}_t(c, () \to e)$$

(RGT)
$$\frac{p; q; 0; v(c) \;\to^* p; q; 0; \mathsf{true}}{p; q; 0;\ \mathsf{trmp}\begin{pmatrix}\mathsf{More}, \\ v_1 :: e_1, \\ v_2 :: e_2\end{pmatrix} \;\to_g p; q; 0; \begin{pmatrix} id \;=\; \mathsf{awaitCl}_t(c,\ \mathsf{this.clfn}(v_1 :: e_1)); \\ \mathsf{trmp}(\mathsf{this.cofn}(v_2 :: id :: e_2)) \end{pmatrix}}$$

(RBT)
$$p; q; 0;\ \mathsf{trmp}\begin{pmatrix}\mathsf{More}, \\ v_1 :: e_1, \\ v_2 :: e_2\end{pmatrix} \;\to_b p; q; 0; \begin{pmatrix} id \;=\; \mathsf{awaitCl}_t(c,\ \mathsf{this.clfn}(v_1 :: e_1)); \\ \mathsf{trmp}(\mathsf{this.cofn}(v_2 :: id :: e_2)) \end{pmatrix}$$

(RT)
$$p; q; 0;\ \mathsf{trmp}(\mathsf{Done},\ v) \;\to p; q; 0;\ v$$

(RG)
$$\frac{p; q; c; e \;\to^* p'; q'; c; v}{p; q; 0;\ \mathsf{awaitCl}_t(c, () \to e) \;\to_g p; q\ \mathsf{msg}(c, v)\ \mathsf{wr}(0, \mathsf{sender}, c); 0;\ v}$$

(RB)
$$p; q; 0;\ \mathsf{awaitCl}_t(c, () \to e) \;\to_b p; q\ \mathsf{msg}(c, v')\ \mathsf{wr}(0, \mathsf{sender}, c); 0;\ v'$$

Figure 14 Evaluation (1/2).

Common Evaluation (Fig. 15). Expressions are reduced under the evaluation context E on the current location (RE), assignment to variables is recorded in the trace (RSET°), accessing a variable is answered by the most recent assignment to it from the trace in the current location (RGET°). For synthetic variables, we use the synthetic store (RGET†, RSET†). Binary operators are defined as unsigned 256 bit integer arithmetic; we only show the rule for addition (ROP). Further, we give rules for conditionals (RT, RF), let binding (RLET) and function application (RLAM) using pattern matching.

Pattern matching (Fig. 13). Matching $[x{\mapsto}v]$ is a partial function, matching patterns x with values v, returning substitution of variables id to values. Matching is recursively defined over pairs; it matches constants to constants, identifiers to values by generating substitutions,

$$
\begin{array}{llll}
(\text{Re}) & p; q; 0;\ E[m] & \rightarrow p'; q'; 0;\ E[m'] & \text{if } p; q; 0; m \ \rightarrow p'; q'; 0; m' \\
(\text{Rget}^{\circ}) & p; q; c;\ \mathsf{this}.i & \rightarrow p; q; c;\ v & \text{if } \mathsf{wr}(c, i, v) \in p \\
(\text{Rget}^{\dagger}) & p; q; c;\ \mathsf{this}.j & \rightarrow p; q; c;\ v & \text{if } \mathsf{wr}(c, j, v) \in q \\
(\text{Rset}^{\circ}) & p; q; c;\ \mathsf{this}.i := v & \rightarrow p\ \mathsf{wr}(c, i, v); q; c;\ () & \\
(\text{Rset}^{\dagger}) & p; q; c;\ \mathsf{this}.j := v & \rightarrow p; q\ \mathsf{wr}(c, j, v); c;\ () & \\
(\text{Rop}) & p; q; c;\ v_0\ +\ v_1 & \rightarrow p; q; c;\ v' & \text{if } v'\ =\ v_0\ +\ v_1 \\
(\text{Rt}) & p; q; c; \begin{pmatrix} \mathsf{if\ let}\ x = v \\ \mathsf{then}\ e_0 \\ \mathsf{else}\ e_1 \end{pmatrix} & \rightarrow p; q; c;\ e_0' & \text{if } e_0'\ =\ [x \Rightarrow v]\ e_0 \\
(\text{Rf}) & p; q; c; \begin{pmatrix} \mathsf{if\ let}\ x = v \\ \mathsf{then}\ e_0 \\ \mathsf{else}\ e_1 \end{pmatrix} & \rightarrow p; q; c;\ e_1 & \\
(\text{Rapp}) & p; q; c;\ (x \rightarrow e)\ v & \rightarrow p; q; c;\ e' & \text{if } e'\ =\ [x \Rightarrow v]\ e \\
(\text{Rlet}) & p; q; c;\ \mathsf{let}\ x = v;\ m & \rightarrow p; q; c;\ m' & \text{if } m'\ =\ [x \Rightarrow v]\ m
\end{array}
$$

Figure 15 Evaluation (2/2).

and fails otherwise. Substitutions $[id \mapsto v]$, in turn, can be applied to terms e, written $[id \mapsto v]\ e$ (capture-avoiding substitution). Substitutions σ compose right-to-left $(\sigma\sigma')x = \sigma(\sigma' x)$.

3.4 Secure Compilation

We prove that the observable behavior of the contract before and after compilation is equivalent. We capture the observable behavior by execution traces and show that trace equivalence holds even when the program is attacked, i.e., reduced by $\rightarrow_b^*$.

Modelling Observable Behavior. The only source of observable nondeterminism in the bad semantics is the evaluation of $\mathsf{awaitCl}_s$ and $\mathsf{awaitCl}_t$. As clients decisions on message sending are influenced by the state of contract variables, tracking incoming client messages and state changes in the trace suffices to capture the observable program behavior. If the observable behavior is the same for the source and the compiled programs, they are indistinguishable. Thus, behavior preservation amounts to trace equality on programs before and after compilation. Further, it suffices to model equality for non-stuck traces. The evaluation gets stuck (program crash) on assertions that guard against deviations from the intended program flow. The Ethereum Virtual Machine reverts contract calls that crash, i.e., state changes of crashed calls do not take effect, hence, stuck traces are not observable.

Since bad evaluation is nondeterministic, we work with not just programs, expressions and configurations, but program sets, expression sets, and configuration sets. Let $p; q; m \Downarrow$ be the trace set of the configuration $p; q; 0; m$, e.g., the set of tuples of the final event-sequence p' and value v of all reduction chains that start in $p; q; 0; m$ and end in $p'; 0; q'; v$. Our trace set definition does not include synthetic events q' of the final configuration. Synthetic events are introduced through compilation; excluding them allows us to put source and target trace sets in relation. Further, let the trace set of a configuration set $T \Downarrow$, be the union of the trace sets for each element:

$$p; q; m \Downarrow\ =\ \{\ (p', v) \mid (p; q; 0; m) \rightarrow_b^* (p'; q'; 0; v)\ \}$$

$$T \Downarrow\ =\ \bigcup_{p;q;m \in T} p; q; m \Downarrow$$

We say that two configuration sets T and S are equivalent, denoted by $T \approx S$, iff T and S have the same traces sets:

$$(T \approx S) \quad \Leftrightarrow \quad (T \Downarrow \;=\; S \Downarrow)$$

By this definition, two expressions that eventually evaluate to the same value with the same trace are related by trace equality. We use this notion of trace equality to prove that a source program is trace-equal to its compiled version by evaluating the compiled program forward $\rightarrow_b^*$ and the original program backward $\leftarrow_b^*$ until configurations converge.

Secure Compilation. Theorem 1 states our correctness property, which says that observable traces generated the malicious evaluation of programs are preserved ($\approx$) by compilation. The malicious evaluation models that client code has been replaced with arbitrary code, while contract code is unchanged. The preservation of observable traces implies the integrity of the (unchanged) contract code. Secure compilation guarantees that developers can write safe programs in the source language without knowledge about the compilation or the distributed execution of client/contract tiers.

▶ **Theorem 1** (Secure Compilation). For each program P over closed terms, the trace set of the program under attack equals the trace set of the compiled program under attack: $\forall P.\, \{\; init_A(comp'(mnf'((P))))\; \} \approx \{\; init_A(P)\; \}$.

We first show that trace equality holds for the different compiler steps. Some compiler steps are defined as a recursive term-to-term transformation on open terms, whereas traceset equality is defined by reducing terms to values, i.e., on closed terms. Since all evaluable programs are closed terms, we show that the compiler steps preserve the traceset of an open term e that is closed by substitution $[x \Mapsto v]$. We formulate the necessary lemmas and sketch the proofs – the detailed proof is in Appendix C.

▶ **Lemma 1** (assoc correct). $\{\, p; q; [x{\Mapsto}v]\, assoc(m)\,\} \approx \{\, p; q; [x{\Mapsto}v]\, m\,\}$

▶ **Lemma 2** (mnf correct). $\{\, p; q; [x{\Mapsto}v]\, mnf(m)\,\} \approx \{\, p; q; [x{\Mapsto}v]\, m\,\}$

▶ **Lemma 3** (mnf' correct). $\{\, init_C(mnf'(d; b; m))\,\} \approx \{\, init_C(d; b; m)\,\}$

▶ **Lemma 4** (comp correct). $\{\, [x{\Mapsto}v]\, init_A(comp(d; b; \mathsf{trmp}(m)))\,\} \approx \{\, init_A(d; b; \mathsf{trmp}([x{\Mapsto}v]\, m))\,\}$

▶ **Lemma 5** (comp' correct). $\{\, [x{\Mapsto}v]\, init_A(comp'(d; b; \mathsf{trmp}(m)))\,\} \approx \{\, init_A(d; b; \mathsf{trmp}([x{\Mapsto}v]\, m))\,\}$

Proof sketch. Lemma 1–5 hold by chain of transitive trace equality relations. We show that a term is trace-equal to the same term after compilation, by evaluating the compiled program ($\rightarrow^*$) and the original program ($\leftarrow^*$) until configurations converge. In the inductive case, we can remove the current compiler step in redex position under traceset equality ($\approx$) since traces before and after applying the compiler step are equal by induction hypothesis.

An interesting case is the proof of $comp$ for $P = d; b; \mathsf{awaitCl}(e_0, () \rightarrow e_1)$. The compiler transforms the remote communication $\mathsf{awaitCl}_s$ into the use of a guard and a trampoline. The compiled program steps to the use of $\mathsf{awaitCl}_t$, the source program to $\mathsf{awaitCl}_s$. In the attacker relation $\rightarrow_b$, arbitrary clients can send arbitrary values with $\mathsf{awaitCl}_t$, leading to additional traces compared to the ones permitted in the source program. where communication is modeled by $\mathsf{awaitCl}_s$. We observe that $\mathsf{awaitCl}_s$ generates the trace elements $\mathsf{msg}(c, v), \mathsf{wr}(0, \text{sender}, c)$ for all v and that $\mathsf{awaitCl}_t$ generates the trace elements $\mathsf{msg}(c', v), \mathsf{wr}(0, \text{sender}, c')$ for all v, c', which differ for $c' \neq c$.

Compilation adds an assert expression (Fig. 10) evaluated after receiving a value from a client. The assert gets stuck for configurations that produce trace elements with $c' \neq c$, removing the traces of such configurations from the trace set, leaving only the traces where $c' = c$. Hence, the trace set before and after compilation is equal under attack.

4 Implementation

Prisma is a language embedded into Scala as a host language. Prisma's features are implemented as a source-to-source macro that transforms Scala code to Scala code. The implementation entails 21 Scala files, 3 412 lines of Scala source code (non-blank, non-comment) licensed under Apache 2.0 Open Source. In addition to the source code of the compiler, we implemented several case studies in Prisma. We further implemented the case studies in Solidity, and test runners in JavaScript for them. A visual comparison of the LOC for the Prisma vs the Solidity case studies can be found in Appendix B.

Prisma contains implementation of the serialisable datatypes, portable between Scala and the EVM (fixed-size arrays, dynamic arrays, unsigned integers of length of powers of two up to 256 bit). Our compiler implementation is based on Scala macros. The compiler phases are macros recursing over the Scala AST: namely the guarding phase, and the "simplifying" phase (including MNF translation, CPS translation of terms, defunctionalisation) and the translation of (a subset) of Scala expressions and types to an custom intermediate representation based on Scala case classes. The intermediate representation is translated to Solidity code, and passed to the Solidity compiler (solc). Our runtime wraps `web3j` [6] (for invoking transactions and interacting with the blockchain in general) and `headlong` [3] (for serialisation/deserialisation in the Ethereum-specific serialisation format), as well code to parse of Solidity and Ethereum error messages to translate them to Scala error messages.

The backend to Solidity is well separated. By disabling the compilation step to Solidity in the compilation pipeline, we could also run distributed code on multiple JVMs instead, for example. In this case, the "contract code" would be executed by a computer to serve as the "server", and other computers would run the "client code" to interact with the server.

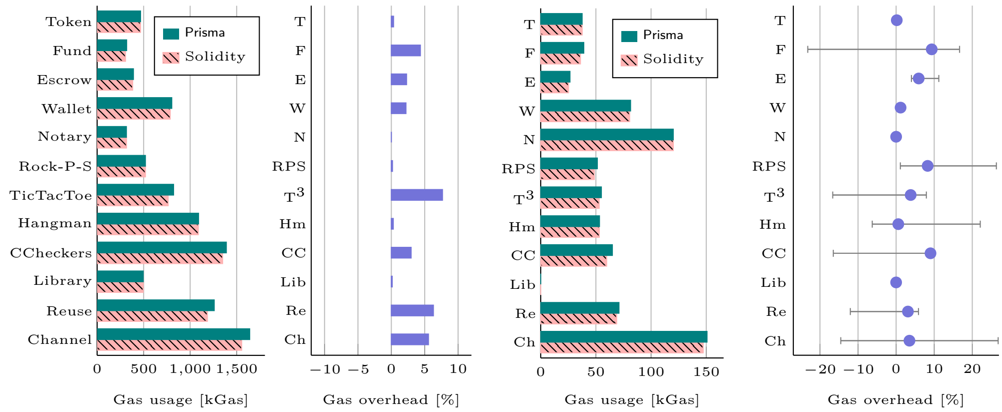
5 Evaluation

We evaluate Prisma along two research questions:

RQ1 *Does Prisma support the most common dApps scenarios?*
RQ2 *Do Prisma's abstractions affect performance?*

Case Studies and Expressiveness (RQ1). Five classes of smart contract applications have been identified [23]: Financial, Wallet, Notary, Game, and Library. To answer RQ1, we implemented at least one case study per category in Prisma. We implemented an ERC-20 Token,[4] a Crowdfunding, and an Escrowing dApp as representatives of financial dApps. We cover *wallets* by implementing a multi-signature wallet, a special type of wallet that provides a transaction voting mechanism by only executing transactions, which are signed by a fixed fraction of the set of owners. We implemented a general-purpose *notary* contract enabling users to store arbitrary data, e.g., document hashes or images, together with a

[4] A study investigating all blocks mined until Sep 15, 2018 [67], found that 72.9 % with an accumulated market capitalization of 12.7 B USD.



(a) Gas usage per deployment. **(b)** Gas overhead per deployment. **(c)** Gas usage per interaction. **(d)** Gas overhead per interaction.

Figure 16 The cost of abstraction. Gas overhead of contracts written with Prisma vs. Solidity. (The right plot displays minima, averages, maxima.)

submission timestamp and the data owner. As *games*, we implemented TicTacToe (Section 2), Rock-Paper-Scissors, Hangman and Chinese Checkers. Rock-Paper-Scissors makes use of timed commitments [19], i.e., all parties commit to a random seed share and open it after all commitments have been posted. The same technique can be used to generate randomness for dApps in a secure way. To reduce expensive code deployment, developers outsource commonly used logic to library contracts. We demonstrate *library*-based development in Prisma by including a TicTacToe library to our case studies and another TicTacToe dApp which uses that library instead of deploying the logic itself.

We implemented a state channel [62, 43, 42] for TicTacToe in Prisma. State channels are *scalability solutions* enabling parties to move their dApp to a non-blockchain consensus system, falling-back to the blockchain in case of disputes.

Our case studies are between 1 K and 7.5 K bytes which is a representative size: Smart contracts are not built for large-scale applications since the gas model limits the maximal computation and storage volumes and causes huge fees for complex applications. The median (average, lower quantile, upper quantile) of the bytecode size of distinct contracts deployed at the time of writing is at 4 K (5.5 K, 1.5 K, 7.5 K) [16]. We further elaborate on the case studies in Appendix A.

Performance of Prisma DApps (RQ2). Performance on the Ethereum blockchain is usually measured in terms of an Ethereum-specific metric called *gas*. Each instruction of the Ethereum Virtual Machine (EVM) consumes gas which needs to be paid for by the users in form of transaction fees credited to the miner. We refer to the Ethereum yellow paper [1] for an overview of the gas consumption of the different EVM instructions. To answer RQ2, we implement our case studies in both Prisma and in Solidity/JavaScript and compare their gas consumption. Unlike prior work, we do not model a custom gas structure, but consider the real EVM gas costs [85].

Experimental setup. We execute each case study on different inputs to achieve different execution patterns that cover all contract functions. Each contract invocation that includes parameters with various sizes (e.g., dynamic length arrays) is executed with a range of realistic inputs, e.g., for Hangman, we consider several words (2 to 40 characters) and different order of guesses, covering games in which the guesser wins and those in which they lose. Prisma and Solidity/JavaScript implementations are executed on the same inputs.

We perform the measurements on a local setup. As the execution in the Ethereum VM is deterministic, a single measurement suffices. We set up the local Ethereum blockchain with *Ganache* (Core v2.13.2) on the latest supported hard fork (Muir Glacier). All contracts are compiled to EVM byte code with *solc* (v0.8.1, optimized on 20 runs). We differentiate contract deployment and contract interaction. Deployment means uploading the contract to the blockchain and initializing its state, which occurs just once per application instance. A single instance typically involves several contract interactions, i.e., transactions calling public contract functions.

Results. Fig. 16 shows the average gas consumption of contract deployment (Fig. 16a) and interaction (Fig. 16c) as well as the relative overhead of Prisma vs. Solidity/JS of deployment (Fig. 16b) and interaction (Fig. 16d). As the gas consumption of contract invocations depends heavily on the executed function, the contract state, and the inputs, we provide the maximal, minimal and averaged overhead. The results show that the average gas consumption of Prisma is close to the one of Solidity/JS. Our compiler achieves a deployment overhead of maximally 6 % The interaction overhead is below 10 % which at most amounts to 3.55 K gas.[5]

Prisma's deployment overhead is mainly due to the automated flow control. To guarantee correct execution, Prisma manages a state variable for dApps with more than one state. The storage reserved for and the code deployed to maintain the state variable cause a constant cost of around 45 K gas. In Solidity, developers manually check whether flow control is needed and, if so, may derive the state from existing contract variables to avoid a state variable if possible.

The Token, Notary, Wallet and Library case studies do not require flow control: each function can be called by any client at any time. Hence, their overhead is small. Escrow, Hangman and Rock-Paper-Scissors require a state variable, also in Solidity – which partially compensates the overhead of Prisma's automated flow control. Crowdfunding, Chinese Checkers, TicTacToe (Library and Channel) do not require an explicit state variable in Solidity, as the state can be derived from the contract variables, e.g., the number of moves. Thus, these case studies have the largest deployment overhead.

While the average relative interaction overhead is constantly below 10 % e.g., in Crowdfunding, TicTacToe Channel, and Rock-Paper-Scissors. Yet, case studies with such spikes also involve interactions that are executed within the same dApp instance with a negative overhead and amortize the costs of more costly transactions. These deviations are also mainly due to automated flow control. In EVM, setting a zero variable to some non-zero value costs more gas (20 K gas) than changing its value (5 K gas) [85], and setting the value to zero saves gas. Occupying and releasing storage via the state variable can cost or save gas in a different way than in traditional dApps without an explicit state variable, leading to different (and even negative) overhead in different transactions.

As Prisma preserves the number of contract interactions, the on-chain execution time of dApps is not affected by our compiler. In Ethereum, the estimated confirmation time for transactions is 3-5 minutes (assuming no congestion), which makes on-chain execution dominate the total execution time. Thus, the overhead of Prisma, if any, is negligible.

Note that per se it is not possible to achieve a better gas consumption in Prisma than in Solidity – every contract compiled from Prisma can be implemented in Solidity. Given the abstractions we offer beyond the traditional development approach, and the sensibility of smart contracts to small changes in instructions, we conclude that our abstractions come

[5] equals 0.59 USD based on gas price and exchange course of April 15, 2021

Table 17 Related work.

Language	Encoding	Perspective	Protocol
Solidity	FSM	Local	Assertions
Obsidian	FSM	Local	Type states
Nomos	MNF	Local	Session types
Prisma	DS	Global	Control flow

with acceptable cost. We are confident that further engineering effort can eliminate the observed overhead.

Threats to validity. The main threat is that the manually written code may be optimized better or worse than the code generated by the compiler. We mitigate this threat by applying all gas optimizations, our compiler performs automatically, to the Solidity implementations. An external threat is that changes in the gas pricing of Ethereum may affect our evaluation. For reproducibility, we state the Ethereum version (hard fork), we used in the paper.

6 Discussion and Related Work

Smart Contract Languages for Enforcing Protocols. We compare Prisma to Solidity, Obsidian [33, 35, 34], and Nomos [39, 40]. We highlight these languages as those also address the correctness of the client–contract interactions. Tab. 17 overviews the features of the surveyed languages for (a) the *perspective* of defining interacting parties, (b) the used *encoding* of the interaction effects, and (c) the method used to check the contract-client interaction *protocol*. Fig. 4, 18, and 19 show code snippets in these languages, each encoding the *TicTacToe* state machine from Fig. 1. All three languages focus solely on the contract and do not state how clients are developed, hence only contract code is shown.

All three approaches take a **local perspective** on interacting parties: Contract and clients are defined separately, and their interaction is encoded by explicit send and receive side effects. In Solidity and Obsidian, receive corresponds to arguments and send to return values of methods defined in the contract classes. In Nomos, send and receive are expressed as procedures operating over a channel – given a channel `c`, sending and receiving is represented by explicit statements (`x = recv c; ...` and `send c x; ...`).

The approaches differ in the **encoding style** of communication effects. Solidity and Obsidian adopt an *FSM-style encoding*: Contract fields encode states, methods encode transitions. The contract in Fig. 4 represents FSM states via the `phase` field with initial state `Funding` (Line 1). The `Fund`, `Move` and `Payout` methods are transitions, e.g., `Payout` transitions the contract into the final state `Closed` (Line 9). The FSM-style encoding results in an implicitly-everywhere concurrent programming model, which is complex to reason about and unfitting for dApps because the execution model of blockchains is inherently sequential – all

```
asset contract TTT {
  state Funding{}; state Executing{}; state Finished{}; state Closed{}
  transaction Fund(TTT@Funding>>(Funding|Executing) this, int c) {
    /*...*/;  if (/* enough funds? */) -> Executing else -> Funding }
  transaction Move(TTT@Executing>>(Executing|Finished) this, int x, int y) {
    /*...*/; if (/* game over? */) -> Finished else -> Executing }
  transaction Payout(TTT@Finished>>Closed this) {
    /*...*/; -> Closed } }
```

Figure 18 Obsidian.

```
1  type Funding   =       int -> +{ notenough: Funding, enough: Executing }
2  type Executing = int -> int -> +{ notdone: Executing, done: Finished }
3  type Finished  =       int -> 1
4  proc contract funding : . |{*}- ($s : Funding) = {
5    a = recv $s; /* ... */
6    if /* enough funds? */ then  $s.notenough; $s <- funding
7                           else  $s.enough; $s <- executing }
8  proc contract executing : . |{*}- ($s : Executing) = {
9    x = recv $s; y = recv $s; /* ... */
10   if /* game over? */ then $s.notdone; $s <- executing
11                       else $s.done; z = recv $s; close $s }
```

Figure 19 Nomos.

$$\textsc{NomosR}\quad \frac{\Psi;\Gamma,(y{:}A)\vdash P :: (c:B)}{\Psi;\Gamma\vdash (y\leftarrow \textsf{recv}\ c;\ \ P) :: (c : A \multimap B)}
\qquad
\textsc{NomosS}\quad \frac{\Psi;\Gamma\vdash P :: (c:B)}{\Psi;\Gamma,(w{:}A)\vdash (\textsf{send}\ c\ w;\ \ P) :: (c : A\otimes B)}$$

$$\textsc{Obsidian}\quad \frac{(\textsf{transaction}\ T\ m(\overline{t.(s \gg s')\ x})\{...\}) \in \text{members}_{t_0}}{\Delta,\overline{e{:}t.s}\vdash e_0.m(\overline{e}) : T \dashv \Delta,\overline{e{:}t.s'}}$$

Figure 20 Excerpts of simplified Nomos and Obsidian typing rules.

method invocations are brought into a world-wide total order. Nomos adopts the *monadic normal form (MNF)* via do-notation to order effects. While the implementation of TicTacToe in FSM style requires three methods(`Fund`, `Move`, `Payout` – one per transition), we only need two methods in MNF-style (`funding`, `executing` – one per state with multiple entry points), and a single method in DS-style (`init`). For instance, the sequence of states and transitions $\mathit{Executing} \xrightarrow{Move(x,y)} \mathit{Finished} \xrightarrow{Payout()} \mathit{Closed}$ in Nomos can be written sequentially in do-notation by inlining the last function which only has a single entry point. Still, do-notation can be cumbersome (e.g., funding and executing in Nomos are separate methods that cannot be inlined since they have multiple entry points and model loops).

All three languages require an **explicit protocol** for governing the send–receive interactions, to ensure that every send effect has a corresponding receive effect in an interacting – separately defined – party. In Solidity, developers express the protocol via run-time assertions to guard against invoking the methods in an incorrect order (e.g., `require(phase==Finished)` in Fig. 4, Line 3). Unlike Solidity, which does not support statically checking protocol compliance, Nomos and Obsidian employ *behavioral typing* for static checks. Deployed contracts may interact with third-party, potentially manipulated clients. Compile-time checking alone cannot provide security guarantees. Yet, complementing run-time enforcement with static checks helps detecting cases that are guaranteed to fail at run time ahead of time.

Obsidian employs typestates to increase safety of contract–client communication. Contracts define a number of typestates; A method call can change the typestate of an object, and calling a method on a receiver that is in the wrong typestate results in a typing error. Each method in Fig. 18 is annotated with the state in which it can be called, e.g., `Payout` requires state `Finished`, and transitions to `Closed` (Line 7).

Nomos employs session types. The session types `Funding`, `Executing`, `Finished` in Fig. 19 encode the protocol. Receiving a message is represented by a function type, e.g., in the `Funding` state, we receive an integer `int ->` ... (Line 1). We respond by either repeating the funding (`Funding`), or continuing to the next state of the protocol (`Executing`). This is represented by internal choice `+{ ... }` that takes multiple possible responses giving each of them a unique label (`notenough` and `enough`). Type `1` indicates the end of a protocol (Line 3). The contract processes `funding` (Line 4) and `executing` (Line 8) implement the protocol. The `recv` operation (Line 5) takes a session-typed channel of form `T -> U`, returns a value of type `T` and changes the type of the channel to `U`. A session type for internal choice (`+{ ... }`), requires the program to select one of the offered labels (e.g., `$s.notenough` in Line 6 and `$s.enough` in Line 7), e.g., in the left and right branch of a conditional statement.

Type systems. We show excerpts of simplified typing rules for Nomos and Obsidian (Fig. 20). Nomos rules have the form $\Psi; \Gamma \vdash P :: (c{:}A)$. A process P offers a channel c of type A with values in context Ψ and channels in Γ. We can see that variables change their type to model the linearity of session types in the NomosS (and NomosR) rule: Sending (and receiving) changes the type of the channel c from $A \multimap B$ to B (and $A \otimes B$ to B). Obsidian rules have the form $\Delta \vdash e{:}t \dashv \Delta'$. An expression e has type t in context Δ and changes Δ to Δ'. We can see that variables change their type on method invocation (Obsidian): A method m in class t_0 with arguments e_i of type t_i, returning T, changes the type state of the arguments from s_i to s_i'. For Prisma, instead, a standard judgement $\Gamma \vdash e : T$ suffices for communication. Variables do not change their type. $\mathsf{awaitCl}(p)\{b\}$ has type T in context Γ if p is a predicate of $Addr$ and b is a pair of $Ether$ and T:

$$\frac{\Gamma \vdash p : Addr \to Bool \qquad \Gamma \vdash b : Ether \times T}{\Gamma \vdash \mathsf{awaitCl}(p)\{\ b\ \} : T}\ \textsc{Prisma}$$

Prisma. As shown in Tab. 17, Prisma occupies an unexplored point in the design space: *global* instead of local perspective on interacting parties, *direct style (DS)* instead of FSM or MNF encoding of effects, and *control flow* instead of extra protocol for governing interactions.

Prisma takes a **global perspective** on interacting parties. The parties execute the same program, where pairs of send and receive actions that "belong together" are encapsulated into a single **direct-style** operation, which is executed differently by sending and receiving parties. Hence, dApps are modeled as sequences and loops of send-receive-instructions shared by interacting parties. Due to the global direct style perspective, it is syntactically impossible to define parties with mismatching send and receive pairs. Hence, a standard System-F-like type system suffices. The interaction protocol follows directly from the sequential **control flow** of interaction points in the program – the compiler can automatically generate access and control guards with correctness guarantees. Semantically, Prisma features a by-default-sequential programming model, intentionally making the sequential execution of methods explicit, including interaction effects.

The global direct-style model also leads to improved design of dApps: No programmatic state management on the contract and no so-called *callback hell* [44] on the client. The direct style is also superior to Nomos' MNF style. The tierless model avoids boilerplate: Client code can directly access public contract variables, unlike JavaScript code, which has to access them via a function call that requires either an await expression or a callback;[6] additionally, the developer has to implement getters for public variables with complex data types such as arrays.[7] We provide some code measurements (lines of code and number of cross-tier control-flow calls) of our Prisma and Solidity/JS dApp case studies in Appendix B.

[6] Obsidian and Nomos do not provide any client design, so we can only compare to Solidity/JavaScript.
[7] For simple data types the getter is generated automatically.

Finally, using one language for both the contract and the clients naturally enables static type safety of values that cross the contract–client boundary: an honest, non-compromized client cannot provide inconsistent input, e.g., with wrong number of parameters or falsely encoded types.[8] In a setting with different language stacks, it is not possible to statically detect type mismatches in the client–contract interaction; e.g.„ Solidity has a type *bytes* for byte arrays, which does not exist in JavaScript (commonly used to implement clients of a Solidity contract). Client developers need to encode byte arrays using hexadecimal string representations starting with “0x”, otherwise they cannot be interpreted by the contract.

Other Related Work. *Smart contract languages.* Harz and Knottenbelt [53] survey smart contract languages, Hu et al. [56] survey smart contract tools and systems, Wöhrer and Zdun [84] give an overview of design patterns in smart contracts. Brünjes and Gabbay [28] distinguish between imperative and functional smart contract programming. *Imperative contracts* are based on the account model; the most prominent language is Solidity [4]. *Functional* ones [29, 77, 76] are based on EUTxO (Extended Unspent Transaction Output) model [50]. State channels [30, 62, 43, 42] optimistically optimize contracts for the functional model. Prisma does not yet support compilation to state channels but we plan to treat them as another kind of tier.

Smart contracts as state machines. Scilla [78] is an automata-based compiler for contracts. FSolidM [60] enables creating contracts via a graphical interface. VeriSolid [61] generates contracts from graphical models enriched with predicates based on computational tree logic. EFSM tools [82] generate contracts from state machines and linear temporal logic. Prisma avoids a separate specification but infers transactions and their order from the control flow of a multitier dApp.

Analysis tools. Durieux et al. [41] and Ferreira et al. [46] empirically validate languages and tools and relate design patterns to security vulnerabilities, extending the survey by Di Angelo and Salzer [20]. Our work is complementary, targeting the correctness of the distributed program flow. For vulnerabilities not related to program flow (e.g., front-running, or bad randomness), developers (using Solidity/JavaScript or Prisma) can use the surveyed analysis tools.

Multitier languages. Multitier programming was pioneered by Hop [79, 80]. Modeling a persistent session in client–server applications with continuations was mentioned by Queinnec [71] and elaborated in Links [37, 49]. Eliom [72] supports bidirectional client–server communication for web applications. ScalaLoci [83] generalizes the multitier model to generic distributed architectures. Our work specializes it to the dApp domain and its specific properties. Giallorenzo et al. [52] establish interesting connections between multitier (subjective) and choreographic (objective) languages – two variants of the global model. Prisma adopts the subjective view, which naturally fits the dApp domain, where a dominant role (contract) controls the execution and diverts control to other parties (clients) to collect their input.

Effectful programs and meta-programming. MNF and CSP are widely discussed as intermediate compiler forms [22, 48, 57, 59, 36]. F# computation expressions [68] support control-flow operators in monadic expressions. OCaml supports a monadic and applicative let [11]: more flexible than do-notation but still restricted to MNF. Idris' !-notation [5] inspired the GHC proposal for monadic inline binding [10]. Scala supports effectful programs through coroutines [8], async/await [73], monadic inline binding [26], Dsl.scala [86] and

[8] Recall that in dApps checking cross-tier type-safety is not a security feature but a design-time safety feature (due to the open-world assumption of the execution model of public ledgers).

a (deprecated) compiler plugin for CPS translation [7]. The dotty-cps-async macro [81] supports async/await and similar effects for the Dotty compiler.

7 Conclusion

We proposed Prisma, the first global language for dApps that features direct style communication. Compared to the state of the art, Prisma (a) enables the implementation of contract and client logic within the same development unit, rendering intricacies of the heterogeneous technology stack obsolete and avoiding boilerplate code, (b) provides support for explicitly encoding the intended program flow and (c) reduces the risk of human failures by enforcing the intended program flow and forcing developers to specify access control.

Unlike previous work that targeted challenges in the development of dApps with advanced typing disciplines e.g., session types, our model does not exhibit visible side effects and gets away with a standard System-F-style type system. We describe the design and the main features of Prisma informally, define its formal semantics, formalize the compilation process and prove it correct. We demonstrate Prisma's applicability via case studies and performance benchmarks.

We plan to generate state channels – to optimistically cost-optimize dApps – similar to how we generate state machines from high-level logic. Further, we believe that our technique for deriving the communication protocol from direct-style control flow generalizes beyond the domain of smart contracts and we will explore its further applicability.

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A Case studies

This section describes the implemented case studies in detail. Bartoletti and Pompianu [23] identify five classes of smart contract applications: Financial, Notary, Game, Wallet, and Library. Our case studies include at least one application per category (Table 21). In addition, we consider scalability solutions.

Financial. These apps include digital tokens, crowdfunding, escrowing, advertisement, insurances and sometimes Ponzi schemes. A study investigating all blocks mined until September 15th, 2018 [67], found that 72.9 % of the high-activity contracts are token contracts compliant to ERC-20 or ERC-721, which have an accumulated market capitalization of US $ 12.7 billion. We have implemented a fungible Prisma token of the ERC-20 standard. Further, we implemented crowdfunding and escrowing case studies. These case studies demonstrate how to send and receive coins with Prisma, which is the basic functionality of financial applications. Other financial use cases can be implemented in Prisma with similar techniques.

Notary. These contracts use the blockchain to store data immutably and persistently, e.g., to certify their ownership. We implemented a general-purpose notary contract enabling users to store arbitrary data, e.g., document hashes or images, together with a submission timestamp and the data owner. This case study demonstrates that Notaries are expressible with Prisma.

Games. We implemented TicTacToe (Section 2), Rock-Paper-Scissors, Hangman and Chinese Checkers. Hangman evolves through multiple phases and hence benefits from the explicit control flow definition in Prisma more than the other game case studies. The game Chinese Checkers is more complex than the others, in regard to the number of parties, the game logic and the number of rounds, and hence, represents larger applications. Rock-Paper-Scissors illustrates how randomness for dApps is securely generated. Every Ethereum transaction, including the executions of contracts, is deterministic – all participants can validate the generation of new blocks. Hence, secure randomness is negotiated among parties: in this case, by making use of timed commitments [19], i.e., all parties commit to a random seed share and open it after all commitments have been posted. The contract uses the sum of all seed shares as randomness. If one party aborts prior to opening its commitment, it is penalized. In Rock-Paper-Scissors both parties commit to their choice – their random share – and open it afterwards. Other games of chance, e.g., gambling contracts, use the same technique.

Wallet. A wallet contract manages digital assets, i.e., cryptocurrencies and tokens, and offers additional features such as shared ownership or daily transaction limits. At August 30, 2019, 3.9 M of 17.9 M (21 %) deployed smart contracts have been different types of wallet contracts [21]. Multi-signature wallets are a special type of wallet that provides a transaction voting mechanism by only executing transactions, which are signed by a fixed fraction of the set of owners. Wallets transfer money and call other contracts in their users stead depending on run-time input, demonstrating calls among contracts in Prisma. Further, a multi-signature wallet uses built-in features of the Ethereum VM for signature validation, i.e., data encoding, hash calculation, and signature verification, showing that these features are supported in Prisma.

Libraries. As the cost of deploying a contract increases with the amount of code in Ethereum, developers try to avoid code repetitions. Contract inheritance does not help: child contracts simply copy the attributes and functions from the parent. Yet, one can outsource commonly used logic to *library contracts* that are deployed once and called by other contracts. For example, the TicTacToe dApp and the TicTacToe channel in our case studies share some logic, e.g., to check the win condition. To demonstrate libraries in Prisma, we include a TicTacToe library to our case studies and another on-chain executed TicTacToe dApp which uses such library instead of deploying the logic itself. Libraries use a call instruction similar to wallets, although the call target is typically known at deployment and can be hard-coded.

Scalability solutions. State channels [62, 43, 42] are scalability solutions, which enable a fixed group of parties to move their dApp to a non-blockchain consensus protocol: the execution falls-back to the blockchain in case of disputes. Similar to multi-signature wallets, state channels use built-in signature validation. We implemented a state channel for TicTacToe[9] to demonstrate that Prisma supports state channels.

B Empirical Evaluation of Design Quality

In Section 6, we argued that with Prisma, (a) we provide communication safety with a standard system-F-like type-system, (b) the program flow can be defined explicitly and is enforced automatically, (c) dApp developers need to master a single technology that covers both tiers, (d) cross-tier type-safety can be checked at compile-time, and (e) the code is simpler and less verbose due to reduced boilerplate code for communication and less control flow jumps. The claims (a), (c), and (d) are a direct consequence of Prisma's design and do not require further evidence. Claim (c) has been formally proven in Section 3. It remains to investigate claim (e), i.e., in which extent Prisma reduces the amount of code and error-prone control-flow jumps.

To this end, we implemented all case studies with equivalent functionality in Prisma and in Solidity/JavaScript. The JavaScript client logic is in direct style using async/await – the Solidity contract needs to be implemented as a finite-state-machine. We keep the client logic

[9] A general solution is a much larger engineering effort and subject of industrial projects [12, 14]

Figure 21 Categories and Cross-tier calls.

Category	Case study	Cross-tier calls	Prisma LoC	Solidity LoC	JavaScript LoC
Financial	Token	4	79	48	50
	Crowdfunding	11	59	27	63
	Escrow	9	63	33	56
Wallet	Multi-signature wallet	3	76	41	52
Notary	General-purpose notary	3	32	16	36
Game	Rock Paper Scissors	12	79	41	77
	TicTacToe	5	61	31	52
	Hangman	15	119	86	83
	Chinese Checkers	4	167	141	47
Library	TicTacToe library	–	167	141	–
	TicTacToe using library	5	53	29	52
Scalability	TicTacToe channel	9	177	56	177

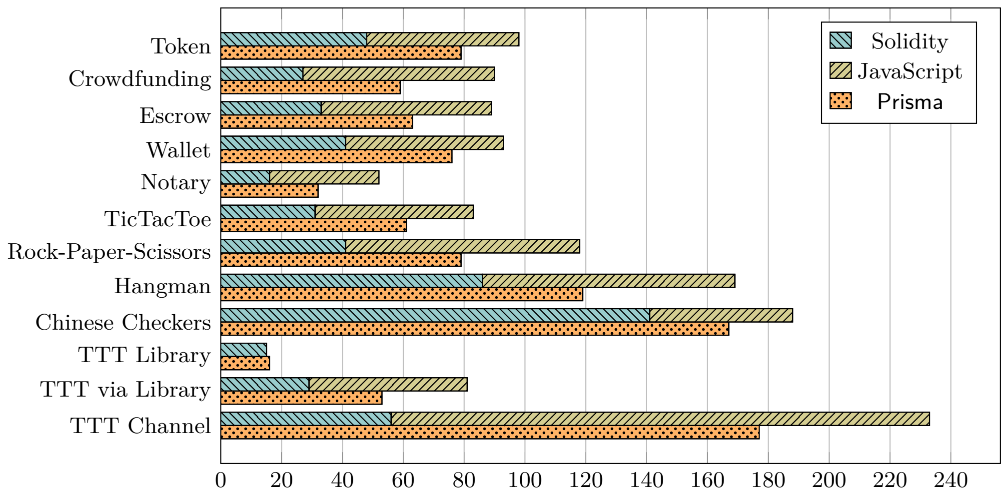

Figure 22 LOC in Solidity/JavaScript and Prisma.

of our case studies (in both, the Prisma and the Solidity implementation) as basic as possible, not to compare the client logic in Scala and in JavaScript but rather focus on the dApp semantics. A complex client logic would shadow the interaction with the contract logic – limited in size due to the gas semantics.

We start with comparing LOCs in the case studies (Figure 22). The results in Figure 22 show that case studies written in Prisma require only 55 – 89 % LOC compared to those implemented in Solidity/JavaScript. One exception is the standalone library, which has no client code and hence does not directly profit from the tierless design.

Second, we consider occurrences of explicit cross-tier control-flow calls in the Solidity/-JavaScript dApps (cf. Table 21), which complicate control flow, compared to Prisma, where cross-tier access is seamless. In the client implementations, 6 – 18 % of all lines trigger a contract interaction passing the control flow to the contract and waiting for the control flow to return. From the contract code in finite-state-machine style, it is not directly apparent at which position the program flow continues, once passed back from clients to contract, i.e., which function is called by the clients next. Direct-style code, on the other hand, ensures that the control flow of the contract always continues in the line that passed the control flow to the client by invoking an `awaitCl` expression.

$$
\begin{array}{lcl}
comp'(d; b; \mathsf{trmp}(m)) & = & d;\ coclfn(b, id, \mathsf{assert(false)}, \mathsf{assert(false)});\ \mathsf{trmp}(comp(m)) \\
& & \text{where} \quad id\ \mathsf{fresh} \\
comp\left(\begin{array}{l} d; coclfn(b, id, \\ \quad e_{1,alt}, \\ \quad e_{2,alt}); \\ tmp \leftarrow_s (() \to e_1);\ e_2 \end{array}\right) & = & \left(\begin{array}{l} d; coclfn(b, id, \\ \quad \mathsf{if\ let}\ (c :: fv(() \to e_1)) = id\ \mathsf{then}\ e_1\ \mathsf{else}\ e'_{1,alt}, \\ \quad \mathsf{if\ let}\ (c :: x :: fv(x \to e'_2)) = id\ \mathsf{then} \\ \qquad \mathsf{assert(this.\,state} == c\ \&\&\ \mathsf{this.\,who(this.\,sender))}; \\ \qquad \mathsf{this.\,state} := 0;\ e'_2 \\ \quad \mathsf{else} \\ \qquad e'_{2,alt}); \\ \mathsf{this.\,who} := e_0;\ \mathsf{this.\,state} := c; \\ (\mathsf{More},\ c :: fv(() \to e_1),\ c :: fv(x \to e'_2)) \end{array}\right) \\
& & \text{where} \quad c\ \mathsf{fresh} \\
& & \text{and} \quad d; coclfn(b, id, e'_{1,alt}, e'_{2,alt}); e'_2 = \\
& & \qquad defun(d; coclfn(b, id, e_{1,alt}, e_{2,alt}); e_2) \\
comp\left(\begin{array}{l} d; coclfn(b, id, e_{1,alt}, e_{2,alt}); \\ x = e_0;\ e_1 \end{array}\right) & = & \left(\begin{array}{l} d; coclfn(b, id, e_{1,alt}, e_{2,alt}); \\ x = e_0;\ defun(e_1) \end{array}\right) \\
comp\left(\begin{array}{l} d; coclfn(b, id, e_{1,alt}, e_{2,alt}); \\ e \end{array}\right) & = & \left(\begin{array}{l} d; coclfn(b, id, e_{1,alt}, e_{2,alt}); \\ e \end{array}\right) \\
coclfn(b, id, e_{1,alt}, e_{2,alt}) & = & (@\mathsf{cl\ this.\,clfn} = id \to e_{1,alt}); (@\mathsf{co\ this.\,cofn} = id \to e_{2,alt}); b
\end{array}
$$

Figure 23 comp′ and comp.

$$
\begin{array}{lcl}
fv(m_0 :: m_1) & = & fv(m_0) \cup fv(m_1) \\
fv(x \to m) & = & fv(m) \setminus fv(x) \\
fv(id) & = & \{id\} \\
fv(m_0\ m_1) & = & fv(m_0) \cup fv(m_1) \\
fv(\mathsf{awaitCl}^*((m_0, () \to m_1)) & = & fv(m_0) \cup fv(m_1) \\
fv(\text{let } x = m_0; m_1) & = & fv(m_0) \cup fv(m_1) \setminus fv(x) \\
fv(\text{this}.i := m_0) & = & fv(m_0) \\
fv(\text{this}.j := m_0) & = & fv(m_0) \\
fv(\text{this}.i) & = & \{\} \\
fv(\text{this}.j) & = & \{\} \\
fv(c) & = & \{\}
\end{array}
$$

Figure 24 Free variables.

C Proofs

We provide the definition of comp′ and comp in Figure 23, the definition for the free variables for a given term fv in Figure 24 and the detailed proofs for the theorem and the lemmas on the following pages.

▶ **Theorem 1** (Secure Compilation). For all programs P over closed terms, the trace set of evaluating the program under attack equals the trace set of evaluating the compiled program under attack, i.e.,

$$\forall P. \quad \{\ init_A(comp'(mnf'((P))))\ \} \approx ... \approx \{\ init_A(P)\ \}$$

Proof.

$$\begin{aligned} & init_A(comp'(mnf'(P))) \\ \overset{\text{Lemma 3}}{\approx} \quad & init_A(mnf'(P)) \\ \overset{\text{Lemma 5}}{\approx} \quad & init_A(P) \end{aligned}$$

◀

Extensions For simplicity, our definition of initialization uses a fixed set of clients. Yet, the malicious semantics does not actually depend on the fixed set of clients, but instead models an attacker that is in control of all clients with the capability of sending messages from any client, not bound to the fixed set. Hence, it is straightforward to extend the proofs to the setting of a dynamic set of clients, e.g., clients joining and leaving at run time.

Further, our trace equality relation defines that all programs in the relation eventually reduce to values, filtering out programs that loop or get stuck. Below, we outline an approach to prove trace equality for looping or stuck programs by showing that such programs loop with the same infinite trace or get stuck at the same trace, respectively. To this end, we track the number of steps done via a step-indexed trace equality relation:

$$p; q; e \Downarrow^n \;=\; \{\ (p', v) \mid (p; q; e) \rightarrow^n (p'; q'; v)\ \} \qquad\qquad T \Downarrow^n = \bigcup_{p;q;e \in T} p; q; e \Downarrow^n$$

With this definition, we can no longer use just equality of traces as the left and right program may take a different number of steps to produce the same events. Instead, we move from an equality relation to a relation stating non-disagreement, which says that – independently of how long we run either statement – the traces will never be in disagreement:

$$(T \approx^n S) \quad \Leftrightarrow \quad (T \Downarrow^n \ \#_{\text{set}}\ S \Downarrow^n)$$

where $\#_{\text{set}}$ is defined on trace sets as

$$T \#_{\text{set}} S \quad \Leftrightarrow \quad (\forall t{\in}T.\ \exists s{\in}S.\ \ t \ \#_{\text{trace}}\ s) \wedge (\forall s{\in}S.\ \exists t{\in}T.\ \ t \ \#_{\text{trace}}\ s)$$

and $\#_{\text{trace}}$ on event traces as

$$\begin{array}{llllll} (ev, & ()) & \#_{\text{trace}} & (ev, & tail_2) & = true \\ (ev, & tail_1) & \#_{\text{trace}} & (ev, & ()) & = true \\ (ev_1, & tail_1) & \#_{\text{trace}} & (ev_2, & tail_2) & = false \\ (ev, & tail_1) & \#_{\text{trace}} & (ev, & tail_2) & = tail_1 \ \#_{\text{trace}}\ tail_2 \end{array} .$$

▶ **Lemma 1** (assoc preserves traces). *assoc* is defined as a recursive term-to-term transformation on open terms, whereas traceset equality is defined by reducing terms to values, i.e., on closed terms. Since all valid programs are closed terms, we show that *assoc* preserves the traceset of an open term e that is closed by substitution $[x \mapsto v]$.

For all terms e, traces p, traces q, values v, patterns x,

$$\{\ p;q;\ [x{\mapsto}v]\ assoc(e)\ \}\ \approx\ ...\ \approx\ \{\ p;q;\ [x{\mapsto}v]\ e\ \}$$

Proof. By induction over term structure.

Case $e\ =\ (\mathsf{let}\ x_1 = (\mathsf{let}\ x_0 = e_0;\ e_1);\ e_2)$.

We know $x_0 \notin fv(e_2)$ since e_2 is not in the scope of the x_0 binding, and that all identifiers are distinct, which can always be achieved by α-renaming.

$$x_0 \notin fv(e_2)$$

According to $\approx$, we only consider terms that reduce to a value. Therefore, let ϕ be the judgement that the term e_0 closed by $[x{\mapsto}v]$ with trace p evaluates to a value v_0 producing trace p_0.

$$\phi\ \equiv\ \ (p;q;\ [x{\mapsto}v]e_0\ \ \rightarrow^*\ \ p\ p_0;q;\ v_0)$$

The lemma holds by the following chain of transitive relations. We evaluate the compiled program from top to bottom ($\rightarrow^*$) and the original program from bottom to top ($\leftarrow^*$) until configurations converge. The induction hypothesis (IH) allows the removal of *assoc* in redex position under traceset equality ($\approx$).

$$\begin{array}{cll}
 & \{\ p;q;\ [x{\mapsto}v]\ assoc(e) & \} \\
\overset{\text{def. } e}{=} & \{\ p;q;\ [x{\mapsto}v]\ assoc(\mathsf{let}\ x_1 = (\mathsf{let}\ x_0 = e_0;\ e_1);\ e_2) & \} \\
\overset{\text{def. } assoc}{=} & \{\ p;q;\ [x{\mapsto}v]\ assoc(\mathsf{let}\ x_0 = e_0;\ assoc(\mathsf{let}\ x_1 = e_1;\ e_2)) & \} \\
\overset{IH}{\approx} & \{\ p;q;\ [x{\mapsto}v]\ \mathsf{let}\ x_0 = e_0;\ assoc(\mathsf{let}\ x_1 = e_1;\ e_2) & \} \\
\overset{\text{def. } \mapsto}{=} & \{\ p;q;\ \mathsf{let}\ x_0 = [x{\mapsto}v]\ e_0;\ [x{\mapsto}v]\ assoc(\mathsf{let}\ x_1 = e_1;\ e_2) & \} \\
\overset{\phi}{\rightarrow^*} & \{\ p\ p_0;q;\ \mathsf{let}\ x_0 = v_0;\ [x{\mapsto}v]\ assoc(\mathsf{let}\ x_1 = e_1;\ e_2)\ \mid\ \forall\ v_0\ p_0,\ \phi & \} \\
\overset{\text{RLET}}{\rightarrow} & \{\ p\ p_0;q;\ [x_0{\mapsto}v_0,\ x{\mapsto}v]\ assoc(\mathsf{let}\ x_1 = e_1;\ e_2)\ \mid\ \forall\ v_0\ p_0,\ \phi & \} \\
\overset{IH}{\approx} & \{\ p\ p_0;q;\ [x_0{\mapsto}v_0,\ x{\mapsto}v]\ \mathsf{let}\ x_1 = e_1;\ e_2\ \mid\ \forall\ v_0\ p_0,\ \phi & \} \\
\overset{\text{def. } \mapsto;\ x_0 \notin fv(e_2)}{=} & \{\ p\ p_0;q;\ \mathsf{let}\ x_1 = [x_0{\mapsto}v_0,\ x{\mapsto}v]e_1;\ [x{\mapsto}v]e_2\ \mid\ \forall\ v_0\ p_0,\ \phi & \} \\
\overset{\text{RLET}}{\leftarrow} & \{\ p\ p_0;q;\ \mathsf{let}\ x_1 = (\mathsf{let}\ x_0 = v_0;\ [x{\mapsto}v]e_1);\ [x{\mapsto}v]e_2\ \mid\ \forall\ v_0\ p_0,\ \phi & \} \\
\overset{\phi}{\leftarrow^*} & \{\ p;q;\ \mathsf{let}\ x_1 = (\mathsf{let}\ x_0 = [x{\mapsto}v]e_0;\ [x{\mapsto}v]e_1);\ [x{\mapsto}v]e_2 & \} \\
\overset{\text{def. } \mapsto}{=} & \{\ p;q;\ [x{\mapsto}v]\ \mathsf{let}\ x_1 = (\mathsf{let}\ x_0 = e_0;\ e_1);\ e_2 & \} \\
\overset{\text{def. } e}{=} & \{\ p;q;\ [x{\mapsto}v]\ e & \}
\end{array}$$

Case $e \neq (\mathsf{let}\ x_1 = (\mathsf{let}\ x_0 = e_0;\ e_1);\ e_2)$.

If e is not of nested let form, we simply apply the definition of *assoc*.

$$\begin{array}{rl} & \{\ p; q;\ [x \mapsto v]\ assoc(e)\ \} \\ \overset{\text{def. } assoc}{=} & \{\ p; q;\ [x \mapsto v]\ e\ \} \end{array}$$

◀

▶ **Lemma 2** (mnf preserves traces). mnf is defined as a recursive term-to-term transformation on open terms, whereas traceset equality is defined by reducing terms to values, i.e., on closed terms. Since all valid programs are closed terms, we show that mnf preserves the traceset of an open term e that is closed by substitution $[x \Mapsto v]$.

For all terms e, traces p, traces q, values v, patterns x,

$$\{\ p; q;\ [x{\Mapsto}v]\ mnf(e)\ \}\ \approx\ ...\ \approx\ \{\ p; q;\ [x{\Mapsto}v]\ e\ \}$$

Proof. By induction over term structure.

Case $e\ =\ e_0\ e_1$.

According to $\approx$, we only consider terms that reduce to a value. Therefore, let ϕ_0 be the judgement that the term e_0 closed by $[x{\Mapsto}v]$ with trace p evaluates to a value v_0 producing trace p_0. Let ϕ_1 be the judgement that the term e_1 closed by $[x{\Mapsto}v]$ with trace $p\ p_0$ evaluates to a value v_1 producing trace $p\ p_0\ p_1$.

$$\begin{aligned}
\phi_0\ &\equiv\quad (p; q;\ [x{\Mapsto}v]\ e_0\quad \to^*\ \ p\ p_0; q;\ v_0)\\
\phi_1\ &\equiv\quad (p\ p_0; q;\ [x{\Mapsto}v]\ e_1\quad \to^*\ \ p\ p_0\ p_1; q;\ v_1)
\end{aligned}$$

Let id_0 be the fresh identifier mnf produces.

$$id_0\ \textsf{fresh}$$

The lemma holds by the following chain of transitive relations. We evaluate the compiled program from top to bottom ($\to^*$) and the original program from bottom to top ($\leftarrow^*$) until configurations converge. The induction hypothesis (IH) allows the removal of mnf in redex position under traceset equality ($\approx$).

$$\begin{array}{rll}
& \{\ p; q;\ [x{\Mapsto}v]\ mnf(e) & \}\\
\stackrel{\text{def. } e}{=} & \{\ p; q;\ [x{\Mapsto}v]\ mnf(e_0\ e_1) & \}\\
\stackrel{\text{def. } mnf}{=} & \{\ p; q;\ [x{\Mapsto}v]\ assoc(\textsf{let}\ id_0 = mnf(e_0);\ assoc(\textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1)) & \}\\
\stackrel{\text{Lemma 1}}{\approx} & \{\ p; q;\ [x{\Mapsto}v]\ \textsf{let}\ id_0 = mnf(e_0);\ assoc(\textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1) & \}\\
\stackrel{\text{def. } \Mapsto}{=} & \{\ p; q;\ \textsf{let}\ id_0 = [x{\Mapsto}v]\ mnf(e_0);\ [x{\Mapsto}v]\ assoc(\textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1) & \}\\
\stackrel{IH}{\approx} & \{\ p; q;\ \textsf{let}\ id_0 = [x{\Mapsto}v]\ e_0;\ [x{\Mapsto}v]\ assoc(\textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1)) & \}\\
\stackrel{\phi_0}{\to^*} & \{\ p\ p_0; q;\ \textsf{let}\ id_0 = v_0;\ [x{\Mapsto}v]\ assoc(\textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1)\mid \forall\ v_0\ p_0,\ \textsf{if}\ \phi_0 & \}\\
\stackrel{\text{RLET}}{\to} & \{\ p\ p_0; q;\ [id_0{\mapsto}v_0,\ x{\Mapsto}v]\ assoc(\textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1)\mid \forall\ v_0\ p_0,\ \textsf{if}\ \phi_0 & \}\\
\stackrel{\text{Lemma 1}}{\approx} & \{\ p\ p_0; q;\ [id_0{\mapsto}v_0,\ x{\Mapsto}v]\ \textsf{let}\ id_1 = mnf(e_1);\ id_0\ id_1\mid \forall\ v_0\ p_0,\ \textsf{if}\ \phi_0 & \}\\
\stackrel{\text{def. } \Mapsto}{=} & \{\ p\ p_0; q;\ \textsf{let}\ id_1 = [id_0{\mapsto}v_0,\ x{\Mapsto}v]\ mnf(e_1);\ v_0\ id_1\mid \forall\ v_0\ p_0,\ \textsf{if}\ \phi_0 & \}\\
\stackrel{IH}{=} & \{\ p\ p_0; q;\ \textsf{let}\ id_1 = [id_0{\mapsto}v_0,\ x{\Mapsto}v]\ e_1;\ v_0\ id_1\mid \forall\ v_0\ p_0,\ \textsf{if}\ \phi_0 & \}\\
\stackrel{id_0\ \textsf{fresh}}{=} & \{\ p\ p_0; q;\ \textsf{let}\ id_1 = [x{\Mapsto}v]\ e_1;\ v_0\ id_1\mid \forall\ v_0\ p_0,\ \textsf{if}\ \phi_0 & \}\\
\stackrel{\phi_1}{\to^*} & \{\ p\ p_0\ p_1; q;\ \textsf{let}\ id_1 = v_1;\ v_0\ id_1\mid \forall\ v_0\ v_1\ p_0\ p_1,\ \textsf{if}\ \phi_0,\ \phi_1 & \}
\end{array}$$

$$\begin{array}{ll}
\overset{\text{RLET}}{\rightarrow} & \{\ p\ p_0\ p_1; q;\ v_0\ v_1 \mid \forall\ v_0\ v_1\ p_0\ p_1,\ \text{if}\ \phi_0,\ \phi_1 \quad \} \\
\overset{\phi_1}{\leftarrow^*} & \{\ p\ p_0; q;\ v_0\ [x{\Mapsto}v]\ e_1 \mid \forall\ v_0\ p_0,\ \text{if}\ \phi_0 \quad \} \\
\overset{\phi_0}{\leftarrow^*} & \{\ p; q;\ ([x{\Mapsto}v]\ e_0)\ [x{\Mapsto}v]\ e_1 \quad \} \\
\overset{\text{def.}\ \Mapsto}{=} & \{\ p; q;\ [x{\Mapsto}v]\ e_0\ e_1 \quad \} \\
\overset{\text{def.}\ e}{=} & \{\ p; q;\ [x{\Mapsto}v]\ e \quad \}
\end{array}$$

Case $e\ =\ \text{let}\ id_0 = e_0;\ e_1.$

According to $\approx$, we only consider terms that reduce to a value. Therefore, let ϕ be the judgement that the term e_0 closed by $[x{\Mapsto}v]$ with trace p evaluates to a value v_0 producing trace p_0.

$$\phi_0 \ \equiv \ (p; q;\ [x{\Mapsto}v]\ e_0 \ \ \rightarrow^* \ p\ p_0; q;\ v_0)$$

The lemma holds by the following chain of transitive relations. We evaluate the compiled program from top to bottom ($\rightarrow^*$) and the original program from bottom to top ($\leftarrow^*$) until configurations converge. The induction hypothesis (IH) allows the removal of mnf in redex position under traceset equality ($\approx$).

$$\begin{array}{ll}
 & \{\ p; q;\ [x{\Mapsto}v]\ mnf(e) \quad \} \\
\overset{\text{def.}\ e}{=} & \{\ p; q;\ [x{\Mapsto}v]\ mnf(\text{let}\ id_0 = v_0;\ e_1) \quad \} \\
\overset{\text{def.}\ mnf}{=} & \{\ p; q;\ [x{\Mapsto}v]\ assoc(\text{let}\ id_0 = mnf(e_0);\ mnf(e_1)) \quad \} \\
\overset{\text{Lemma 1}}{\approx} & \{\ p; q;\ [x{\Mapsto}v]\ \text{let}\ id_0 = mnf(e_0);\ mnf(e_1) \quad \} \\
\overset{\text{def.}\ \Mapsto}{=} & \{\ p; q;\ \text{let}\ id_0 = [x{\Mapsto}v]\ mnf(e_0);\ [x{\Mapsto}v]\ mnf(e_1) \quad \} \\
\overset{IH}{\approx} & \{\ p; q;\ \text{let}\ id_0 = [x{\Mapsto}v]\ e_0;\ [x{\Mapsto}v]\ mnf(e_1) \quad \} \\
\overset{\phi_0}{\rightarrow^*} & \{\ p\ p_0; q;\ \text{let}\ id_0 = v_0;\ [x{\Mapsto}v]\ mnf(e_1) \mid \forall\ v_0\ p_0,\ \text{if}\ \phi_0 \ \} \\
\overset{\text{RLET}}{\rightarrow} & \{\ p\ p_0; q;\ [id_0{\mapsto}v_0,\ x{\Mapsto}v]\ mnf(e_1) \mid \forall\ v_0\ p_0,\ \text{if}\ \phi_0 \quad \} \\
\overset{IH}{\approx} & \{\ p\ p_0; q;\ [id_0{\mapsto}v_0,\ x{\Mapsto}v]\ e_1 \mid \forall\ v_0\ p_0,\ \text{if}\ \phi_0 \quad \} \\
\overset{\text{RLET}}{\leftarrow} & \{\ p\ p_0; q;\ [x{\Mapsto}v]\ \text{let}\ id_0 = v_0;\ e_1 \mid \forall\ v_0\ p_0,\ \text{if}\ \phi_0 \quad \} \\
\overset{\phi_0}{\leftarrow^*} & \{\ p; q;\ [x{\Mapsto}v]\ \text{let}\ id_0 = e_0;\ e_1 \quad \} \\
\overset{\text{def.}\ e}{=} & \{\ p; q;\ [x{\Mapsto}v]\ e \quad \}
\end{array}$$

Case The other cases of e are proved analogously.

◀

▶ **Lemma 3** (mnf' preserves trace)**.** mnf' is defined on programs. To evaluate a program, it is initialized with a set of clients A. mnf' preserves the traceset of (closed) programs P for any set of clients A.

For all P,

$$\{\ init_A(mnf'(P))\ \} \ \approx\ ...\ \approx\ \{\ init_A(P)\ \}$$

Proof. By induction over term structure.

Case $P \ = \ (d; b;\ e_0)$.

Initializing the definitions $d; b$ with A produces the trace p and the state q.

$$init_A(d; b) \ = \ p; q$$

According to $\approx$, we only consider terms that reduce to a value. Therefore, let ϕ be the judgement that the term e_0 closed by $[x{\mapsto}v]$ in trace p produces a value v_0 and trace p_0.

$$\phi \ \equiv \quad (p; q;\ e_0 \quad \rightarrow^* \ \ p\ p_0; q;\ v_0)$$

The lemma holds by the following chain of transitive relations. We evaluate the compiled program from top to bottom ($\rightarrow^*$) and the original program from bottom to top ($\leftarrow^*$) until configurations converge, using Lemma 2.

$$\begin{array}{rll}
 & \{\ init_A(mnf'(P)) & \} \\
\overset{\text{def. } P}{=} & \{\ init_A(mnf'(d; b;\ e_0)) & \} \\
\overset{\text{def. } mnf'}{=} & \{\ init_A(d; b;\ \mathsf{trmp}(mnfe(\mathsf{Done}(e_0)))) & \} \\
\overset{\text{def. } init_A}{=} & \{\ p; q;\ \mathsf{trmp}(mnfe(\mathsf{Done}(e_0))) & \} \\
\overset{\text{Lemma 2}}{\approx} & \{\ p; q;\ \mathsf{trmp}(\mathsf{Done}(e_0)) & \} \\
\overset{\phi}{\rightarrow^*} & \{\ p\ p_0; q;\ \mathsf{trmp}(\mathsf{Done}(v_0)) \mid \forall\ v_0\ p_0,\ \mathsf{if}\ \phi & \} \\
\overset{\text{RDONE}}{\rightarrow} & \{\ p\ p_0; q;\ v_0 \mid \forall\ v_0\ p_0,\ \mathsf{if}\ \phi & \} \\
\overset{\phi}{\leftarrow^*} & \{\ p; q;\ e_0 & \} \\
\overset{\text{def. } init_A}{=} & \{\ init_A(d; b;\ e_0) & \} \\
\overset{\text{def. } P}{=} & \{\ init_A(P) & \}
\end{array}$$

◀

▶ **Lemma 4** (comp preserves traces). *comp* is defined on programs. To evaluate a program, it is initialized with a set of clients A. *comp* preserves the traceset of (closed) programs P for any set of clients A.

For all definitions b, definitions d, terms e, values v, patterns x,

$$\{\ [x \Rrightarrow v]\ init_A(comp(d; b;\ \mathsf{trmp}(e)))\ \} \ \approx\ \dots\ \approx\ \{\ init_A(d; b;\ \mathsf{trmp}([x \Rrightarrow v]\ e))\ \}$$

Proof. By induction over term structure.

Case $\quad e \ = \ \mathsf{let}\ x = \mathsf{awaitCl}_s((e_0,\ () \rightarrow e_1));\ e_2.$

comp expects the definitions b to be of form:

$$b \ = \ \begin{pmatrix} @\mathsf{cl\ this.clfn} \ = \ id \rightarrow e_{1,alt}; \\ @\mathsf{co\ this.cofn} \ = \ id \rightarrow e_{2,alt}; \\ b_{rest} \end{pmatrix}$$

comp is defined recursively and applied to the term e_2. Intuitively, *comp* transforms e_2 to e_2' and b to b' by moving the part of e_2 that comes after the $\mathsf{awaitCl}_s$ call into the cofn definition inside b. The recursive call is given as follows:

$$(d; b';\ \mathsf{trmp}(e_2')) \ = \ comp(d; b;\ \mathsf{trmp}(e_2))$$

$$b' \ = \ \begin{pmatrix} @\mathsf{cl\ this.clfn} \ = \ id \rightarrow e'_{1,alt}; \\ @\mathsf{co\ this.cofn} \ = \ id \rightarrow e'_{2,alt}; \\ b_{rest} \end{pmatrix}$$

After the recursive call, *comp* moves the transformed e_2' into the cofn definition, resulting in e' and b'' with $e''_{1,alt}$ and $e''_{2,alt}$.

$$\phi \ \equiv \ \Big(\{\ d; b'';\ \mathsf{trmp}(e')\ \Big\} \ = \ \Big\{\ comp(d; b;\ \mathsf{trmp}(e))\ \}\ \Big)$$

$$b'' \ = \ \begin{pmatrix} @\mathsf{cl\ this.clfn} \ = \ id \rightarrow e''_{1,alt}; \\ @\mathsf{co\ this.cofn} \ = \ id \rightarrow e''_{2,alt}; \\ b_{rest} \end{pmatrix}$$

$$e''_{1,alt} \ = \ \begin{pmatrix} \mathsf{if\ let}\ (c :: fv(() \rightarrow e_1)) \ = \ id \\ \mathsf{then}\ e_1 \\ \mathsf{else}\ e'_{2,alt} \end{pmatrix}$$

$$e''_{2,alt} \ = \ \begin{pmatrix} \mathsf{if\ let}\ (c :: x :: fv(x \rightarrow e_2')) \ = \ id \\ \mathsf{then\ assert}(\mathsf{this.state} == c\ \&\&\ \mathsf{this.sender} == \mathsf{this.who}); \\ \qquad \mathsf{this.state} := 0;\ e_2' \\ \mathsf{else}\ e'_{2,alt} \end{pmatrix}$$

Let $p; q$ be the trace and state produced by initializing $d; b$ with A, and $p; q'$ for initializing $d; b'$, and $p; q''$ for initializing $d; b''$.

$$init_A(d; b) \ = \ p; q$$

$$init_A(d; b') \ = \ p; q'$$

$$init_A(d; b'') \ = \ p; q''$$

According to $\approx$, we only consider terms that reduce to a value. Therefore, let ϕ_0 be the judgement that the term e_0 closed by $[x{\Rightarrow}v]$ in trace p produces a value v_0 and trace p_1.

$$\phi_0(q_\phi) \;=\; (p; q_\phi;\; [x{\Rightarrow}v]\; e_0 \;\; \rightarrow p\; p_1; q_\phi;\; v_0)$$

We define ϕ_1 based on ϕ:

$$
\begin{array}{l}
\;\phi \\
= \\
\;\left\{\; d; b'';\; \mathsf{trmp}(e') \;\right\} \;=\; \left\{\; comp(d; b;\; \mathsf{trmp}(e)) \;\right\} \\
\;\rightarrow generalize\; [x{\Rightarrow}v]\; init_A(...) \\
\;\left\{\; [x{\Rightarrow}v]\; init_A(d; b'';\; \mathsf{trmp}(e')) \;\right\} \;=\; \left\{\; [x{\Rightarrow}v]\; init_A(comp(d; b;\; \mathsf{trmp}(e))) \;\right\} \\
\;\rightarrow (= \;\; \rightarrow \;\approx) \\
\;\left\{\; [x{\Rightarrow}v]\; init_A(d; b'';\; \mathsf{trmp}(e')) \;\right\} \;\approx\; \left\{\; [x{\Rightarrow}v]\; init_A(comp(d; b;\; \mathsf{trmp}(e))) \;\right\} \\
\;\rightarrow IH \\
\;\left\{\; [x{\Rightarrow}v]\; init_A(d; b'';\; \mathsf{trmp}(e')) \;\right\} \;\approx\; \left\{\; init_A(d; b;\; \mathsf{trmp}([x{\Rightarrow}v]\; e)) \;\right\} \\
\;\rightarrow \text{def. } init_A \\
\;\left\{\; p; q'';\; \mathsf{trmp}([x{\Rightarrow}v]\; e') \;\right\} \;\approx\; \left\{\; p; q;\; \mathsf{trmp}([x{\Rightarrow}v]\; e) \;\right\} \\
\equiv \\
\;\phi_1
\end{array}
$$

The lemma holds by the following chain of transitive relations. We evaluate the compiled program from top to bottom ($\rightarrow^*$) and the original program from bottom to top ($\leftarrow^*$) until configurations converge.

$$
\begin{array}{cl}
 & \left\{\; [x{\Rightarrow}v]\; init_A(comp(d; b;\; \mathsf{trmp}(e))) \;\right\} \\
\overset{\text{def. } e}{=} & \left\{\; [x{\Rightarrow}v]\; init_A(comp(d; b;\; \mathsf{trmp}(\mathsf{let}\; x_3 = \mathsf{awaitCl}_s((e_0,\; () \rightarrow e_1));\; e_2)) \;\right\} \\
\overset{\text{def. } comp}{=} & \left\{ \begin{array}{l} [x{\Rightarrow}v]\; init_A(d; b'';\; \mathsf{trmp}(\\ \mathsf{this.who} := e_0;\; \mathsf{this.state} := c; \\ \mathsf{More}(c :: \; fv(() \rightarrow e_1),\; c ::\; fv(x \rightarrow e_2')))) \end{array} \right\} \\
\overset{\text{def. } init_A}{=} & \left\{ \begin{array}{l} p; q'';\; [x{\Rightarrow}v]\; \mathsf{trmp}(\\ \mathsf{this.who} := e_0;\; \mathsf{this.state} := c; \\ \mathsf{More}(c :: \; fv(() \rightarrow e_1),\; c ::\; fv(x \rightarrow e_2')) \end{array} \right\} \\
\overset{\text{def. } \Rightarrow}{=} & \left\{ \begin{array}{l} p; q'';\; \mathsf{trmp}(\\ \mathsf{this.who} := [x{\Rightarrow}v]\; e_0;\; \mathsf{this.state} := c; \\ \mathsf{More}(c :: [x{\Rightarrow}v]\; fv(() \rightarrow e_1),\; c :: [x{\Rightarrow}v]\; fv(x \rightarrow e_2'))) \end{array} \right\} \\
\overset{\phi_0(q'')}{\rightarrow^*} & \left\{ \begin{array}{l} p\; p_1;\; q'';\; \mathsf{trmp}(\\ \mathsf{this.who} := v_0;\; \mathsf{this.state} := c; \\ \mathsf{More}(c :: [x{\Rightarrow}v]\; fv(() \rightarrow e_1),\; c :: [x{\Rightarrow}v]\; fv(x \rightarrow e_2'))) \\ \mid \forall\; v_0\; p_1,\; \mathsf{if}\; \phi_0 \end{array} \right\} \\
\overset{\text{RSET}\dagger,\; \text{RSET}\dagger}{\rightarrow} & \left\{ \begin{array}{l} p\; p_1;\; q''\; [\mathsf{who} \mapsto v_0,\; \mathsf{state} \mapsto c]; \\ \mathsf{trmp}(\mathsf{More}(c :: [x{\Rightarrow}v]\; fv(() \rightarrow e_1)),\; c :: [x{\Rightarrow}v]\; fv(x \rightarrow e_2')) \\ \mid \forall\; v_0\; p_1,\; \mathsf{if}\; \phi_0 \end{array} \right\}
\end{array}
$$

$$
\begin{aligned}
&\overset{\textsc{RMore}}{\rightarrow} && \left\{ \begin{array}{l} p\ p_1;\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ tmp\ \leftarrow_t \mathsf{this.clfn}(c :: [x \Mapsto v]\ \mathit{fv}(() \rightarrow e_1)); \\ \mathsf{trmp}(\mathsf{this.cofn}(c :: tmp :: [x \Mapsto v]\ \mathit{fv}(x \rightarrow e_2'))) \\ \mid \forall\ v_0\ p_1,\ \text{if}\ \phi_0 \end{array} \right\} \\
&\overset{\textsc{RBt}}{\rightarrow} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ \text{let}\ tmp = v_2';\ \mathsf{trmp}(\mathsf{this.cofn}(c :: tmp :: [x \Mapsto v]\ \mathit{fv}(x \rightarrow e_2'))) \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ \phi_0 \end{array} \right\} \\
&\overset{case\ v_0' = v_0}{=} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ \text{let}\ tmp = v_2';\ \mathsf{trmp}(\mathsf{this.cofn}(c :: tmp :: [x \Mapsto v]\ \mathit{fv}(x \rightarrow e_2'))) \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ v_0' \neq v_0,\ \phi_0 \\ \hline p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ \text{let}\ tmp = v_2';\ \mathsf{trmp}(\mathsf{this.cofn}(c :: tmp :: [x \Mapsto v]\ \mathit{fv}(x \rightarrow e_2'))) \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ v_0' = v_0,\ \phi_0 \end{array} \right\} \\
&\overset{\textsc{RLet},\ \textsc{RGet},\ \textsc{RApp},\ \textsc{Rt},\ \textsc{RGet},\ \textsc{ROp},\ \textsc{RGet},\ \textsc{RGet},\ \textsc{ROp},\ \textsc{ROp}}{\rightarrow^*} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ \mathsf{trmp}(\mathsf{assert}(\mathsf{false});\ \mathsf{this.state} := 0;\ [x \Mapsto v_2',\ x \Mapsto v]\ e_2')) \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ v_0' \neq v_0,\ \phi_0 \\ \hline p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ \mathsf{trmp}(\mathsf{assert}(\mathsf{true});\ \mathsf{this.state} := 0;\ [x \Mapsto v_2',\ x \Mapsto v]\ e_2') \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ v_0' = v_0,\ \phi_0 \end{array} \right\} \\
&\overset{\text{def.}\ \approx}{\approx} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto c]; \\ \mathsf{trmp}(\mathsf{assert}(\mathsf{true});\ \mathsf{this.state} := 0;\ [x \Mapsto v_2',\ x \Mapsto v]\ e_2') \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ v_0' = v_0,\ \phi_0 \end{array} \right\} \\
&\overset{\textsc{RLet},\ \textsc{RSet}}{\rightarrow} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0', v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0');\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto 0]; \\ \mathsf{trmp}([x \Mapsto v_2',\ x \Mapsto v]\ e_2') \\ \mid \forall\ v_0\ p_1\ v_0'\ v_2',\ \text{if}\ v_0' = v_0,\ \phi_0 \end{array} \right\} \\
&\overset{v_0' = v_0}{=} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0, v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0);\ q''\ [\mathsf{who}\mapsto v_0,\ \mathsf{state}\mapsto 0]; \\ \mathsf{trmp}([x \Mapsto v_2',\ x \Mapsto v]\ e_2') \\ \mid \forall\ v_0\ p_1\ v_2',\ \text{if}\ \phi_0 \end{array} \right\} \\
&\overset{\phi_1}{\approx} && \left\{ \begin{array}{l} p\ p_1\ \mathsf{msg}(v_0, v_2')\ \mathsf{wr}(0, \mathsf{sender}, v_0);\ q; \\ \mathsf{trmp}([x \Mapsto v_2',\ x \Mapsto v]\ e_2) \\ \mid \forall\ v_0\ p_1\ v_2',\ \text{if}\ \phi_0 \end{array} \right\} \\
&\overset{\textsc{RLet},\ \textsc{RBs}}{\leftarrow} && \left\{ \begin{array}{l} p\ p_1;\ q;\ \mathsf{trmp}(\text{let}\ x = \mathsf{awaitCl}_s(v_0,\ () \rightarrow [x \Mapsto v]\ e_1);\ [x \Mapsto v]\ e_2) \\ \mid \forall\ v_0\ p_1,\ \text{if}\ \phi_0 \end{array} \right\} \\
&\overset{\phi_0(q)}{\leftarrow^*} && \left\{\ p; q;\ \mathsf{trmp}(\text{let}\ x = \mathsf{awaitCl}_s([x \Mapsto v]\ e_0,\ () \rightarrow [x \Mapsto v]\ e_1);\ [x \Mapsto v]\ e_2)\ \right\} \\
&\overset{\text{def.}\ \Mapsto}{=} && \left\{\ p; q;\ \mathsf{trmp}([x \Mapsto v]\ \text{let}\ x = \mathsf{awaitCl}_s(e_0,\ () \rightarrow e_1);\ e_2)\ \right\} \\
&\overset{\text{def.}\ e}{=} && \left\{\ p; q;\ \mathsf{trmp}([x \Mapsto v]\ e)\ \right\} \\
&\overset{\text{def.}\ init_A}{=} && \left\{\ init_A(b; d;\ \mathsf{trmp}([x \Mapsto v]\ e))\ \right\}
\end{aligned}
$$

Case $e = x_0$.

Let $p; q$ be the trace and state produced by initializing $d; b$ with A.

$$init_A(d; b) \ = \ p; q$$

The traceset equality holds by definition of *comp* and $init_A$.

$$
\begin{array}{rll}
& \{ & [x \mapsto v]\ init_A(comp(d; b;\ \mathsf{trmp}(e))) \quad \} \\
\overset{\text{def. } e}{=} & \{ & [x \mapsto v]\ init_A(comp(d; b;\ \mathsf{trmp}(x_0))) \quad \} \\
\overset{\text{def. } comp}{=} & \{ & [x \mapsto v]\ init_A(d; b;\ comp(\mathsf{trmp}(x_0))) \quad \} \\
\overset{\text{def. } init_A}{=} & \{ & p; q;\ [x \mapsto v]\ \mathsf{trmp}(x_0) \quad \} \\
\overset{\text{def. } \mapsto}{=} & \{ & p; q;\ \mathsf{trmp}([x \mapsto v]\ x_0) \quad \} \\
\overset{\text{def. } e}{=} & \{ & p; q;\ \mathsf{trmp}([x \mapsto v]\ e) \quad \} \\
\overset{\text{def. } init_A}{=} & \{ & init_A(d; b;\ \mathsf{trmp}([x \mapsto v]\ e)) \quad \}
\end{array}
$$

◀

▶ **Lemma 5** (comp' preserves traces)**.** $comp'$ is defined on programs. To evaluate a program, it is initialized with a set of clients A. $comp'$ preserves the traceset of (closed) programs P for any set of clients A.

For all definitions b, definitions d, terms e_0,

$$\{\ init_A(comp'(d; b;\ \mathsf{trmp}(e_0)))\ \} \ \approx\ \ldots\ \approx\ \{\ init_A(d; b';\ \mathsf{trmp}(e_0))\ \}$$

Proof. By induction over term structure.

Case $\quad P\ =\ (d; b;\ e_0)$.

Intuitively, $comp'$ prepends the definitions b with initial definitions for clfn and cofn that only contain assert(false), such that $comp$ can be applied.

$$b'\ =\ \left(\begin{array}{l} @\mathsf{cl}\ \mathsf{this.clfn}\ =\ id \to \mathsf{assert(false)};\\ @\mathsf{co}\ \mathsf{this.cofn}\ =\ id \to \mathsf{assert(false)};\\ b \end{array}\right)$$

The lemma holds by definition of $comp'$, and Lemma 4.

$$\begin{array}{rl} & \{\ init_A(comp'(d; b;\ \mathsf{trmp}(e_0)))\ \}\\ \overset{\text{def. } comp'}{=} & \{\ init_A(comp(d; b';\ \mathsf{trmp}(e_0))\ \}\\ \overset{\text{Lemma 4}}{\approx} & \{\ init_A(d; b';\ \mathsf{trmp}(e_0))\ \} \end{array}$$

◀