

A Novel Approach to Parameterized verification of Cache Coherence Protocols

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Abstract—Parameterized verification of parameterized protocols like cache coherence protocols is important but hard. Our tool paraVerifier handles this hard problem in a unified framework: (1) it automatically discovers auxiliary invariants and the corresponding causal relations from a small reference instance of the verified protocol; (2) the above invariants and causal relation information are automatically generalized into a parameterized form to construct a parameterized formal proof in a theorem prover (e.g., Isabelle). The principle underlying the generalization is the symmetry mapping. Our method is successfully applied to typical benchmarks including snoopy-based and directory-based benchmarks. Another novel feature of our method lies in that the final verification result of a protocol is provided by a formal and readable proof.

I. INTRODUCTION

Verification of parameterized concurrent systems is interesting in the area of formal methods, mainly due to the practical importance of such systems. Parameterized systems exist in many important application areas, including cache coherence, security, and network communication protocols. The hardness of parameterized verification is mainly due to the requirement of correctness that the desired properties should hold in any instance of the parameterized system. The model checkers, although powerful in verification of non-parameterized systems, become impractical to verify parameterized systems, as they can verify only an instance of the parameterized system in each execution. A desirable approach is to provide a proof that the correctness holds for any instance.

Related Work: There have been a lot of studies in the field of parameterized verification [1], [2], [3], [4], [5], [6], [7], [8], [9]. Among them, the ‘invisible invariants’ method [3] is an automatic technique for parameterized verification. In this method, auxiliary invariants are computed in a finite system instance to aid inductive invariant checking. The CMP method [6] adopts parameter abstraction and guard strengthening to verify a safety property *inv* of a parameterized system. An abstract instance of the parameterized protocol is constructed by a counter-example-guided refinement process in an informal way.

The degree of scalability and automation are two critical aspects of approaches to parameterized verification. The verification of real-world parameterized systems is, however, rarely scalable and automatic. For instance, up to now, the verification of a real-world benchmark FLASH requires human guidance in the existing successful verifications [10], [11], [6]. In order to effectively verify complex parameterized protocols like FLASH, there are two critical problems. The first one is how to find a set of sufficient and necessary invariants without

(or with less) human intervention, which is a core problem in this field. The second one is the rigorousness of the verification. The theory foundation of a parameterized verification technique and its soundness are only discussed in a paper proof style in previous work. It is preferable to formulate all the verification in a publicly-recognized trust-worthy framework like a theorem prover [6]. However, theorem proving in a theorem prover like Isabelle is interactive, not automatic.

In order to solve the parameterized verification in a both automatic and rigorous way, we design a tool called paraVerifier, which is based on a simple but elegant theory. Three kinds of causal relations are introduced, which are essentially special cases of the general induction rule. Then, a so-called consistency lemma is proposed, which is the cornerstone of our method. Notably, the theory foundation itself is verified as a formal theory in Isabelle, which is the formal library for verifying protocol case studies. The library provides basic types and constant definitions to model protocol cases and lemmas to prove properties.

Our tool paraVerifier is composed of two parts: an invariant finder *invFinder* and a proof generator *proofGen*. Given a protocol \mathcal{P} and a property *inv*, *invFinder* tries to find useful auxiliary invariants and causal relations which are capable of proving *inv*. To construct auxiliary invariants and causal relations, we employ heuristics inspired by consistency relation. Also, when several candidate invariants are obtained using the heuristics, we use oracles such as a model checker and an SMT-solver to check each of them under a small reference model of \mathcal{P} , and chooses the one that has been verified.

After *invFinder* finds the auxiliary invariants and causal relations, *proofGen* generalizes them into parameterized forms, which are then used to construct a completely parameterized formal proof in a theorem prover (e.g., Isabelle) to model \mathcal{P} and to prove the property *inv*. After the base theory is imported, the generated proof is checked automatically. Usually, a proof is done interactively. Special efforts in the design of the proof generation are made in order to make the proof checking automatic.

II. PRELIMINARIES

There are three kinds of *variables*: 1) simple identifier, denoted by a string; 2) element of an array, denoted by a string followed by a natural inside a square bracket. E.g., *arr*[*i*] indicates the *i*th element of the array *arr*; 3) field of a record, denoted by a string followed by a dot and then another string. E.g., *rcd.f* indicates the field *f* of the record *rcd*.

Each variable is associated with its *type*, which can be an enumeration, natural number, and Boolean.

Expressions and *formulas* are defined mutually recursively. *Expressions* can be simple or compound. A simple expression is either a variable or a constant while a compound expression is constructed with the *ite*(if-then-else) form $f ? e_1 : e_2$, where e_1 and e_2 are expressions, and f is a formula. A *formula* can be an atomic formula or a compound formula. An atomic formula can be a boolean variable or constant, or in the equivalence form $e_1 \doteq e_2$, where e_1 and e_2 are two expressions. A *formula* can also be constructed by using the logic connectives, including negation (!), conjunction (\wedge), disjunction (\vee), implication (\rightarrow).

An *assignment* is a mapping from a variable to an expression, and is denoted with the assigning operation symbol “:=”. A *statement* α is a set of assignments which are executed in parallel, e.g., $x_1 := e_1; x_2 := e_2; \dots; x_k := e_k$. If an assignment maps a variable to a (constant) value, then we say it is a *value-assignment*. We use $\alpha|_x$ to denote the expression assigned to x under the statement α . For example, let α be $\{arr[1] := C; x := false\}$, then $\alpha|_x$ returns *false*. A *state* is an instantaneous snapshot of its behavior given by a set of value-assignments.

For every expression e and formula f , we denote the value of e (or f) under an state $s :: var \Rightarrow valueType$ as $\mathbb{A}[e, s]$ (or $\mathbb{B}[f, s]$). For the state s and a formula f , we write $s \models f$ to mean $\mathbb{B}[f, s] = true$. Formal semantics of expressions and formulas are given in HOL as usual, which is shown in [12].

For an expression e and a statement $\alpha = x_1 := e_1; x_2 := e_2; \dots; x_k := e_k$, we use $vars(\alpha)$ to denote the variables to be assigned $\{x_1, x_2, \dots, x_k\}$; and use e^α to denote the expression transformed from e by substituting each x_i with e_i simultaneously. Similarly, for a formula f and a statement $\alpha = x_1 := e_1; x_2 := e_2; \dots; x_k := e_k$, we use f^α to denote the formula transformed from f by substituting each x_i with e_i . Moreover, f^α can be regarded as the weakest precondition of formula f w.r.t. statement α , and we denote $preCond(f, \alpha) \equiv f^\alpha$. Noting that a state transition is caused by an execution of the statement, formally, we define: $s \xrightarrow{\alpha} s' \equiv (\forall x \in vars(\alpha). s'(x) = \mathbb{A}[\alpha|_x, s]) \wedge (\forall x \notin vars(\alpha). s'(x) = s(x))$.

A *rule* r is a pair $\langle g, \alpha \rangle$, where g is a formula and is called the *guard* of rule r , and α is a statement and is called the *action* of rule r . For convenience, we denote a rule with the guard g and the statement α as $g \triangleright \alpha$. Also, we denote $act(g \triangleright \alpha) \equiv \alpha$ and $pre(g \triangleright \alpha) \equiv g$. If the guard g is satisfied at the state s , then α can be executed, thus, a new state s' is derived, and we say the rule $g \triangleright \alpha$ is triggered at s , and transited to s' . Formally, we define: $s \xrightarrow{r} s' \equiv s \models pre(r) \wedge s \xrightarrow{act(r)} s'$.

A *protocol* \mathcal{P} is a pair (I, R) , where I is a set of *formulas* and is called the initializing formula set, and R is a set of rules. As usual, the reachable state set of protocol $\mathcal{P} = (I, R)$, denoted as $reachableSet(\mathcal{P})$, can be defined inductively: (1) a state s is in $reachableSet(\mathcal{P})$ if there exists a formula $f \in I$,

and $s \models f$; (2) a state s is in $reachableSet(\mathcal{P})$ if there exist a state s_0 and a rule $r \in R$ such that $s_0 \in reachableSet(\mathcal{P})$ and $s_0 \xrightarrow{r} s$.

A parameterized object(T) is simple a function from a natural number to T, namely of type $nat \Rightarrow T$. For instance, a parameterized formula pf is of type $nat \Rightarrow formula$, and we define $forallForm(1, pf) \equiv pf(1)$, and $forallForm((n+1), pf) \equiv forallForm(n, pf) \wedge pf(n+1)$. $existsForm(1, pf) \equiv pf(1)$, and $existsForm((n+1), pf) \equiv existsForm(n, pf) \vee pf(n+1)$.

Now we use a simple example to illustrate the above definitions by a simple mutual exclusion protocol with N nodes. Let I, T, C, and E be three enumerating values, x, n are simple and array variables, N a natural number, $pini(N)$ the predicate to specify the inial state, $prules(N)$ a HOL-notation to denote a set of the four rules of the protocol, $mutualInv(i, j)$ a property that $n[i]$ and $n[j]$ cannot be C at the same time. We want to verify that $mutualInv(i, j)$ holds for any $i \leq N$, $j \leq N$ s.t. $i \neq j$. **example 1 has been thoroughly checked**

Example 1 *Mutual-exclusion example.*

```
assignN(i) ≡ n[i] ≡ I
pini(N) ≡ x ≠ true ∧ forallForm(N, assignN)
try(i) ≡ n[i] ≡ I ▷ n[i] := T
crit(i) ≡ n[i] ≡ T ∧ x ≠ true ▷ n[i] := C; x := false
exit(i) ≡ n[i] ≡ C ▷ n[i] := E
idle(i) ≡ n[i] ≡ E ▷ n[i] := I; x := true
prules(N) ≡ {r. ∃ i. i ≤ N ∧ (r = crit(i) ∨ r = exit(i) ∨ r = idle(i) ∨ r = try(i))}
mutualEx(N) ≡ (pIni(N), prules(N))
mutualInv(i, j) ≡ ! (n[i] ≡ C ∧ n[j] ≡ C)
```

III. CAUSAL RELATIONS AND CONSISTENCY LEMMA

A novel feature of our work lies in that three kinds of causal relations are exploited, which capture whether and how the execution of a particular protocol rule changes the protocol state variables appearing in an invariant. Consider a rule r , a formula f , and a formula set fs , three kinds of causal relations are defined as follows:

Definition 1 We define the following relations:
 $invHoldRule_1 :: state \times formula \times rule \Rightarrow bool$,
 $invHoldRule_2 :: state \times formula \times rule \Rightarrow bool$,
 $invHoldRule_3 :: state \times formula \times rule \times ruleset \Rightarrow bool$, and
 $invHoldRule_3 :: state \times formula \times rule \times ruleset \Rightarrow bool$.

- 1) $invHoldRule_1(s, f, r) \equiv s \models pre(r) \rightarrow s \models preCond(f, act(r));^1$
- 2) $invHoldRule_2(s, f, r) \equiv s \models f \leftrightarrow s \models preCond(f, act(r));$
- 3) $invHoldRule_3(s, f, r, fs) \equiv \exists f' \in fs. (s \models (f' \wedge pre(r)) \rightarrow s \models preCond(f, act(r)));$
- 4) $invHoldRule(s, f, r, fs) \equiv s \models invHoldRule_1(s, f, r) \vee s \models invHoldRule_2(s, f, r) \vee s \models invHoldRule_3(s, f, r, fs).$

The relation $invHoldRule(s, f, r, fs)$ defines a causality relation between f , r , and fs , which guarantees that if each

¹Here \rightarrow and \leftrightarrow are HOL connectives. Throughout this work, we use HOL as our meta-logic, and embed our protocol description in HOL including descriptions of rules and properties.

formula in fs holds before the execution of rule r , then f holds after the execution of rule r . This includes three cases. 1) $\text{invHoldRule}_1(s, f, r)$ means that after rule r is executed, f becomes true immediately; 2) $\text{invHoldRule}_2(s, f, r)$ states that $\text{preCond}(S, f)$ is equivalent to f , which intuitively means that none of the state variables in f is changed, and the execution of statement S does not affect the evaluation of f ; 3) $\text{invHoldRule}_3(s, f, r, fs)$ states that there exist another invariant $f' \in fs$ such that the conjunction of the guard of r and f' implies the precondition $\text{preCond}(S, f)$. We can also view $\text{invHoldRule}_{1,3}$ as three special kind of inductive tactics, which can be applied to prove each formula in fs holds at each inductive protocol rule cases. Note that the three kinds of inductive tactics can be done by a theorem prover, which is the cornerstone of our work. Only after the theorem prover is told which one among the three kinds of tactics is to be used, it can prove automatically. Without the fine-grained classification, the theorem prover can't solve the proof goals. In the procedure of automatic proof generation, `proofGen` generates proof scripts which contain enough application of the three kinds of tactics and guide the theorem prover to finish the proof.

With the invHoldRule relation, we define a consistency relation $\text{consistent}(invs, inis, rs)$ between a protocol $(inis, rs)$ and a set of invariants $invs = \{inv_1, \dots, inv_n\}$.

Definition 2 We define a relation $\text{consistent} :: \text{formula set} \times \text{formula set} \times \text{rule set} \Rightarrow \text{bool}$. $\text{consistent}(invs, inis, rs)$ holds if the following conditions hold:

- 1) for all formulas $inv \in invs$ and $ini \in inis$ and all states s , $s \models ini$ implies $s \models inv$;
- 2) for all formulas $inv \in invs$ and rules $r \in rs$ and all states s , $\text{invHoldRule}(s, inv, r, invs)$

Example 2 Let us define a set of auxiliary invariants:

```

invOnXC(i)  $\equiv$   $!(x \doteq \text{true} \wedge n[i] \doteq C)$ 
invOnXE(i)  $\equiv$   $(x \doteq \text{true} \wedge n[i] \doteq E)$ 
aux1(i, j)  $\equiv$   $!(n[i] \doteq C \wedge n[j] \doteq E)$ 
aux2(i, j)  $\equiv$   $!(n[i] \doteq E \wedge n[j] \doteq E)$ 
pinvs(N)  $\equiv$   $\{f. \exists iInv1\ iInv2. iInv1 \leq N \wedge iInv2 \leq N \wedge$ 
 $iInv1 \neq iInv2 \wedge f = \text{mutualInv } iInv1\ iInv2\}$ 
 $\vee (\exists iInv1. iInv1 \leq N \wedge f = \text{invOnXC } iInv1)$ 
 $\vee (\exists iInv1. iInv1 \leq N \wedge f = \text{invOnXE } iInv1)$ 
 $\vee (\exists iInv1\ iInv2. iInv1 \leq N \wedge iInv2 \leq N \wedge iInv1 \neq iInv2$ 
 $\wedge f = \text{aux1 } iInv1\ iInv2)$ 
 $\vee (\exists iInv1\ iInv2. iInv1 \leq N \wedge iInv2 \leq N \wedge iInv1 \neq iInv2$ 
 $\wedge f = \text{aux2 } iInv1\ iInv2)\}$ .

```

In the following discussion, we assume that $inv = \text{mutual}(i_1, i_2)$, $r = \text{crit}(iR_1)$, $rs = \text{pinvs}(N)$, and assumptions $i_1 \neq N$, $i_2 \neq N$, $i_1 \neq i_2$, and $iR_1 \leq N$ hold.

- $\text{invHoldRule}_2(s, inv, r)$, where $i_1 \neq iR_1$, and $i_2 \neq iR_1$, since $\text{preCond}(\text{act}(r), inv) = inv$.
- $\text{invHoldRule}_3(s, inv, r, invs)$, where $i_1 = iR_1$. Since $\text{invOnXC}(i_2) \in invs$, $\text{preCond}(\text{act}(r), inv) = (C \doteq C \wedge n[i_2] \doteq C)$, and $s \models (\text{invOnXC}(i_2) \wedge \text{pre}(\text{crit}(iR_1)))$ implies $s \models (C \doteq C \wedge n[i_2] \doteq C)$.

- $\text{invHoldRule}_3(s, inv, r, invs)$, where $i_2 = iR_1$. Since $\text{invOnXC}(i_1) \in invs$, $\text{preCond}(\text{act}(r), inv) = (n[i_1] \doteq C \wedge C \doteq C)$, and $s \models (\text{invOnXC}(i_2) \wedge \text{pre}(\text{crit}(iR_1)))$ implies $s \models (n[i_1] \doteq C \wedge C \doteq C)$.

For any invariant $inv \in invs$, inv holds at a reachable state s of a protocol $P = (ini, rs)$ if the consistency relation $\text{consistent}(invs, inis, rs)$ holds. The following lemma formalizes the essence of the aforementioned causal relation, and is called consistency lemma.

Theorem 1 If $P = (ini, rs)$, $\text{consistent}(invs, ini, rs)$, and $s \in \text{reachableSet}(P)$, then for all inv s.t. $inv \in invs$, $s \models inv$.

Theorem 1 is our main weapon to prove. Let us recall the proof goal set in 1: the mutual exclusion property holds for each reachable state of the mutual-exclusion protocol. In order to prove the goal, we prove a more general result:

Lemma 2 If $P = (\text{pini}(N), \text{prules}(N))$ is the protocol listed in example 1, $s \in \text{reachableSet}(P)$, and $0 < N$, and $\text{pinvs}(N)$ is the set of formulas in example 2, then for any inv s.t. $inv \in \text{pinvs}(N)$, $s \models inv$.

Proof: By theorem 1, we only need to prove that parts (1) and (2) of the relation $\text{consistent}(\text{pinvs}(N), \text{pini}(N), \text{prules}(N))$ hold. Part (1) can be checked routinely. Part (2) can be proved by case analysis on a formula $f \in invs$ and a rule $r \in rs$. Example 2 has checked one case: $f = \text{mutual}(i_1, i_2)$, $r = \text{crit}(iR_1)$. Other cases can be analyzed similarly. ■

In order to apply the consistency lemma to prove that a given property inv (e.g., the mutual exclusion property) holds for each reachable state of a protocol $P = (inis, rs)$ (e.g., mutual-exclusion protocol), we need to solve two problems. First, we need to construct a set of auxiliary invariants $invs$ which contains inv and satisfies $\text{consistent}(invs, inis, rs)$. By applying the consistency lemma, we decompose the original problem of invariant checking into that of checking the causal relation between some $f \in invs$ and $r \in rs$. The latter needs case analysis on the form of f and r . Only if a proof script contains sufficient information on the case splitting and the kind of causal relation to be checked in each subcase, Isabelle can help us to automatically check it. How to generate automatically such a proof, which can be run in Isabelle, is the second problem.

Our solutions to the two problems are as follows: Given a protocol, `invFinder` finds all the necessary ground auxiliary invariants from a small instance of the protocol in `Murphi`. This step solves the first problem. A table `protocol.tbl` is worked out to store the set of ground invariants and causal relations, which are then used by `proofGen` to create an Isabelle proof script which models and verifies the protocol in a parameterized form. In this step, ground invariants are generalized into a parameterized form, and accordingly ground

causal relations are adopted to create parameterized proof commands which essentially proves the existence of the parameterized causal relations. This solves the second problem. At last, the Isabelle proof script is fed into Isabelle to check the correctness of the protocol.

IV. SEARCHING AUXILIARY INVARIANTS

Algorithm 1: Algorithm: *invFinder*

Input: Initially given invariants F , a protocol $\mathcal{P} = \langle I, R \rangle$

Output: A set of tuples which represent causal relations between concrete rules and invariants:

```

1  $A \leftarrow F$ ;
2  $tuples \leftarrow []$ ;
3  $newInvs \leftarrow F$ ;
4 while  $newInvs$  is not empty do
5    $f \leftarrow newInvs.dequeue$ ;
6   for  $r \in R$  do
7      $paras \leftarrow \text{Policy}(r, f)$ ;
8     for  $para \in paras$  do
9        $cr \leftarrow \text{apply}(r, para)$ ;
10       $newInvOpt, rel \leftarrow \text{coreFinder}(cr, f, A)$ ;
11       $tuples \leftarrow tuples @ \langle r, para, f, rel \rangle$ ;
12      if  $newInvOpt \neq \text{NONE}$  then
13         $newInv \leftarrow \text{get}(newInvOpt)$ ;
14         $newInvs.enqueue(newInv)$ ;
15         $A \leftarrow A \cup \{newInv\}$ ;
16 return  $tuples$ ;

```

Given a protocol \mathcal{P} and a property set F containing invariant formulas we want to verify, *invFinder* aims to find useful auxiliary invariants and causal relations which are capable of proving any element in F . A set A is used to store all the invariants found up to now, and is initialized as F . A queue $newInvs$ is used to store new invariants which have not been checked, and is initialized as F . A relation table $tuples$ is used to record the causal relation between a parameterized rule in some parameter setting and a concrete invariant. Initially, $tuples$ are set as NULL. *invFinder* works iteratively in a semi-proving and semi-searching way. In each iteration, the head element f of $newInvs$ is popped, then $\text{Policy}(r, f)$ generates groups of parameters $paras$ according to r and f by some policy. For each parameter $para$ in $paras$, it is applied to instantiate r into a concrete rule cr . Here $\text{apply}(r, para) = r$ if r contains no array-variables and $para = []$; otherwise $\text{apply}(r, para) = r(para_{[1]}, \dots, para_{[|para|]})$. Then $\text{coreFinder}(cr, f, A)$ is called to check whether a causal relation exists between cr and f ; if there is such one relation item, the relation item rel and a formula option $newInvOpt$ is returned; otherwise a run-time error occurs in coreFinder , which indicates no proof can be found. In the first case, a tuple $\langle r, para, f, rel \rangle$ will be inserted into $tuples$; If the formula option $newInvOpt$ is NONE, then no new invariant formula is generated; otherwise $newInvOpt = \text{Some}(f')$ for some formula f' , then $\text{get}(newInvOpt)$ returns f' , and the new invariant formula f' will be pushed into the queue $newInvs$ and inserted into the invariant set A . The above searching

process is executed until $newInvs$ becomes empty. At last, the table $tuples$ is returned.

Here we still use the mutual exclusion protocol to illustrate the main ideas of *invFinder*. Let $P = (pini(N), prules(N))$ is the protocol listed in example 1, $f = \text{mutualInv}(1, 2)$, and $F = \{f\}$. The output of Algorithm 1 is to construct useful auxiliary invariants in example 2 and causal relations used in Lemma 2. By this example, the parameter generation policy *Policy* and the core invariant searching function *coreFinder* will be illustrated in Section IV-A and IV-B.

A. Parameter Generation Policy

Let $r = \text{crit}(i)$, which is a parameterized rule. How many groups of rule parameters are needed to instantiate r into concrete rules in order to compute more auxiliary invariants and causal relations between these concrete rules and f . For instance, [1], [2], and [3] are three groups to instantiate r into $\text{crit}(1)$, $\text{crit}(2)$, and $\text{crit}(3)$. However, a natural question is: is it enough for our parameter generation policy to generate only these three groups of concrete rules. Do we need another group parameter [4] to instantiate r ? Roughly speaking, after the generation of concrete rules according to the policy, enough auxiliary invariants and causal relations should be computed to generate a proof as shown in Lemma 2. In detail, through the computation of $\text{coreFinder}(cr, f, A)$ by using $\text{crit}(1)$, $\text{crit}(2)$ and $\text{crit}(3)$ with f , adopting the information generated from the generated auxiliary invariants and causal relations should derive a proof of case on f and crit in Example 2 which also involves three subcases. Here [4] is not necessary because [3] and [4] are “equivalent” by our *Policy*.

In order to formulate the main ideas of our parameter generation policy, we introduce the concept of permutation modulo to symmetry relation \simeq_m^n , and a quotient set of perms_m^n (the set of all n -permutations of m) under the relation. Here an n -permutation of m is ordered arrangement of an n -element subset of an m -element set $I = \{i. 0 < i \leq m\}$. We use a list xs with size n to stand for an n -permutation of m . For instance, [1, 2] is a 2-permutation of 3. $xs[i]$ and $|xs|$ denote the i -th element and the length of xs respectively. If $xs[i] = i$ for all $i \leq |xs|$, we call it identical permutation.

Definition 3 Let m and n be two natural numbers, where $n \leq m$, L and L' are two lists which stand for two n -permutations of m ,

- 1) $L \simeq_m^n L' \equiv (|L| = |L'| = n) \wedge (\forall i. i < |L| \wedge L[i] \leq m - n \longrightarrow L'[i] = L'[i])$.
- 2) $L \simeq_m^n L' \equiv L \sim_m^n L' \wedge L' \sim_m^n L$.
- 3) $\text{semiP}(m, n, S) \equiv (\forall L \in \text{perms}_m^n \exists L' \in S. L \simeq_m^n L') \wedge (\forall L \in S. \forall L' \in S. L \neq L' \longrightarrow \neg(L \simeq_m^n L'))$.
- 4) A set S is called a quotient of the set perms_m^n under the relation \simeq_m^n if $\text{semiP}(m, n, S)$.

The definition of relation \simeq_m^n (item 1 and 2 in Definition 3) directly leads to the following lemma.

Lemma 3 If $L \simeq_{m+n}^n L'$, then for any $0 < i \leq |L|$, any $0 < j \leq m$, $L_{[i]} = j$ if and only if $L'_{[i]} = j$.

For instance, let $L = [2, 3]$ and $L' = [2, 4]$, then $L \simeq_4^2 L'$. Due to Lemma 3, we can analyze a group of concrete parameters by analyzing only one of them as a representative. Keeping this in mind, let us look at the following lemma, which together with Lemma 3 is the theoretical basis of our policy.

Lemma 4 Let S be a set s.t. $\text{semiP}(m, n, S)$,

- 1) for any $L \in \text{perms}_m^n$, there exists a $L' \in S$ s.t. $L \simeq_m^n L'$.
- 2) let $L \in S$, $L' \in S$, if $L \neq L'$, then there exists two indices $i \leq m$ and $j \leq n$ such that $L_{[i]} = j$ and $L'_{[i]} \neq j$.

Lemma 4 shows 1) completeness of S w.r.t. the set perms_m^n under the relation \simeq , 2) the distinction between two different elements in S . Therefore, S has covered all analysing patterns according to the aforementioned comparing scheme between elements of L with numbers $j < n - m$. Moreover, the case patterns represented by different elements in S are different from each other. This fact can be illustrated by the following example.

Example 3 Let $m = 3$, $n = 1$, $S = \{[1], [2], [3]\}$ and $\text{semiP}(m, n, S)$, let LR be an element in S , there are three cases:

- 1) $LR = [1]$: it is a special case where $LR_{[1]} = 1$;
- 2) $LR = [2]$: it is a special case where $LR_{[1]} = 2$;
- 3) $LR = [3]$: it is a special case where $LR_{[1]} \neq 1$ and $LR_{[1]} \neq 2$;
- 4) $[3] \simeq_3^1 [4]$, they are both the special cases where $LR_{[1]} \neq 1$ and $LR_{[1]} \neq 2$.

Note that the above cases are mutually disjoint and their disjunction is true.

In Algorithm 1, a concrete formula cinv is popped from the queue newInvs , which can be seen as a normalized instantiation of some parameterized formula pinv .

Definition 4 A concrete invariant formula cinv is normalized w.r.t a parameterized invariant pinv if there exists no array variable in cinv and $\text{pinv} = \text{cinv}$ or there exists an identical permutation LI with $|LI| > 0$ such that $\text{cinv} = \text{pinv}(1, \dots, |LI|)$;

Any normalized cinv containing array variables is obtained by instantiating a parameterized invariant pinv with a parameter list which is an identical permutation LI (i.e., the j^{th} parameter is j itself $LI_{[j]} = j$). For instance, the concrete formula $!(n[1] \doteq C \wedge n[2] \doteq C)$ is obtained by instantiating $\text{mutualInv}(i_1, i_2)$ with $[1, 2]$. Thus, consider a list of parameter LR which is used to instantiate a parameterized rule pr , we have $LR_{[i]} = j$ (or $LR_{[i]} \neq j$) is equivalent to

$LR_{[i]} = LI_{[j]}$ (or $LR_{[i]} \neq LI_{[j]}$), which is a factor to specify a case by comparing $LR_{[i]}$ with $LI_{[j]}$. This will be the key idea used to generalize the concrete invariants and causal relations computed by invFinder .

Let cinv be a normalized concrete invariant w.r.t. a parameterized invariant pinv , pr be a parameterized rule, m be the number of actual parameters occurring in cinv , and n be the number of formal parameters occurring in pr , our policy is to compute a quotient of perms_m^n , denoted as $\text{cmpSemiperm}(m+n, n)$, and use elements of it as a group of parameters to instantiate pr into a set crs of concrete rules.² For instance, for the invariant $!(n[1] \doteq C \wedge n[2] \doteq C)$ (or $\text{mutualInv}(1, 2)$), three groups of parameters $[1]$, $[2]$, $[3]$ are used to instantiate crit respectively, each of the instantiation results will be used to check which kind of causal relation exists. We will explain the checking work done by coreFinder , which is illustrated in the following subsection.

B. Core Searching Algorithm

For a cinv and a rule $r \in \text{crs}$, the core part of the invFinder tool is shown in Algorithm 2. It needs to call two oracles. The first one, denoted by chk , checks whether a ground formula is an invariant. Such an oracle can be implemented by translating the formula into a formula in SMV, and calling SMV to check whether it is an invariant in a given small reference model of the protocol. If the reference model is too small to check the invariant, then the formula will be checked by Murphi in a big reference model. The second oracle, denoted by tautChk , checks whether a formula is a tautology. Such a tautology checker is implemented by translating the formula into a form in the SMT (SAT Modulo Theories) format, and checking it by an SMT solver such as Z3.

Algorithm 2: Core Searching Algorithm: coreFinder

Input: $r, \text{inv}, \text{invs}$
Output: A formula option f , a new causal relation rel

```

1  $g \leftarrow$  the guard of  $r$ ,  $S \leftarrow$  the statement of  $r$ ;
2  $\text{inv}' \leftarrow \text{preCond}(\text{inv}, S)$ ;
3 if  $\text{inv} = \text{inv}'$  then
4    $\text{relItem} \leftarrow (r, \text{inv}, \text{invRule}_2, -)$ ;
5   return  $(\text{NONE}, \text{relItem})$ ;
6 else if  $\text{tautChk}(g \rightarrow \text{inv}') = \text{true}$  then
7    $\text{relItem} \leftarrow (r, \text{inv}, \text{invRule}_1, -)$ ;
8   return  $(\text{NONE}, \text{relItem})$ ;
9 else
10   $\text{candidates} \leftarrow \text{subsets}(\text{decompose}(\text{dualNeg}(\text{inv}') \wedge g))$ ;
11   $\text{newInv} \leftarrow \text{choose}(\text{chk}, \text{candidates})$ ;
12   $\text{relItem} \leftarrow (r, \text{inv}, \text{invRule}_3, \text{newInv})$ ;
13  if  $\text{isNew}(\text{newInv}, \text{invs})$  then
14     $\text{newInv} \leftarrow \text{normalize}(\text{newInv})$ ;
15    return  $(\text{SOME}(\text{newInv}), \text{relItem})$ ;
16  else
17    return  $(\text{NONE}, \text{relItem})$ ;
```

²the details of computing $\text{cmpSemiperm}(m+n, n)$ can be found in [12].

Input parameters of Algorithm 2 include a rule instance r , an invariant inv , a sets of invariants $invs$. The sets $invs$ stores the auxiliary invariants constructed up to now. The algorithm searches for new invariants and constructs the causal relation between the rule instance r and the invariant inv . The algorithm returns a formula option and a causal relation item between r and inv . A formula option value NONE indicates that no new invariant is found while $SOME(f)$ indicates a new auxiliary invariant f is searched.

Algorithm *coreFinder* works as follows: after computing the pre-condition inv' (line 2), which is the weakest precondition of the input formula inv w.r.t. S , the algorithm takes further operations according to the cases it faces with:

- (1) If $inv = inv'$, meaning that statement S does not change inv , then no new invariant is created, and new causal relation item marked with tag $invHoldRule_2$ is recorded between r and inv .
- (2) If $tautChk$ verifies that $g \dashv\vdash inv'$ is a tautology, then no new invariant is created, and the new causal relation item marked with tag $invHoldRule_1$ is recorded between r and inv .
- (3) If neither of the above two cases holds, then a new auxiliary invariant $newInv$ will be constructed, which will make the causal relation $invHoldRule_3$ to hold. The candidate set is $subsets(decompose(dualNeg(inv') \bar{\wedge} g))$, where $decompose(f)$ decompose f into a set of subformulas f_i such that each f_i is not of a conjunction form and f is semantically equivalent to $f_1 \bar{\wedge} f_2 \bar{\wedge} \dots \bar{\wedge} f_N$. $dualNeg(!f)$ returns f . $subsets(S)$ denotes the power set of S . A proper formula is chosen from the candidate set to construct a new invariant $newInv$. This is accomplished by the `choose` function, which calls the oracle `chk` to verify whether a formula is an invariant in the given reference model. After $newInv$ is chosen, the function *isNew* checks whether this invariant is new w.r.t. $newInvs$ or $invs$. If this is the case, the invariant will be normalized, and then be added into $newInvs$, and the new causal relation item marked with tag $invRule_3$ will be added into the causal relations. The meaning of the word “new” is modulo to the symmetry relation. E.g., $mutualInv(1, 2)$ is equivalent to $mutualInv(2, 1)$ in a symmetry view.

TABLE I
A FRAGMENT OF OUTPUT OF *invFinder*

rule	ruleParas	inv	causal relation	f'
..
crit	[1]	$mutualInv(1,2)$	$invHoldRule3$	$invOnXC(2)$
crit	[2]	$mutualInv(1,2)$	$invHoldRule3$	$invOnXC(1)$
crit	[3]	$mutualInv(1,2)$	$invHoldRule2$	

Let us continue the example in the end of subsection IV-A. After the three iterations of computations of *coreFinder* on $crit(1)$, $crit(2)$, $crit(3)$ and $mutualInv(1, 2)$, the according

output of the *invFinder*, which is stored in file `mutual.tbl`, is shown in Table I. In the table, each line records the index of a normalized invariant, name of a parameterized rule, the rule parameters to instantiate the rule, a causal relation between the ground invariant and a kind of causal relation which involves the kind and proper formulas f' in need (which are used to construct causal relations $invHoldRule_3$). $mutualInv(1, 2)$.

V. GENERALIZATION

Intuitively, generalization means that a concrete index (formula or rule) is generalized into a set of concrete indices (formulas or rules), which can be formalized by a symbolic index (formula or rules) with side conditions specified by constraint formulas. In order to do this, we adopt a new constructor to model symbolic index or symbolic value $symb(str)$, where str is a string. We use N to denote $symb("N")$, which formalizes the size of a parameterized protocol instance. A concrete index i can be transformed into a symbolic one by some special strategy g , namely $symbolize(g, i) = symb(g(i))$. In this work, two special transforming function $flnv(i) = "iInv" \hat{\wedge} itoa(i)$ and $flr(i) = "iR" \hat{\wedge} itoa(i)$, where $itoa(i)$ is the standard function transforming an integer i into a string. We use special symbols $iInv_i$ to denote $symbolize(fInv, i)$; and iR_i to denote $symbolize(fIr, i)$. The former formalizes a symbolic parameter of a parameterized formula, and the latter a symbolic parameter of a parameterized rule. Accordingly, we define $symbolize2f(g, inv)$ (or $symbolize2r(g, r)$), which returns the symbolic transformation result to a concrete formula inv (or rule r) by replacing a concrete index i occurring in inv (or r) with a symbolic index $symbolize(g, i)$.

There are two main kinds of generalization in our work: (1) generalization of a normalized invariant into a symbolic one. The resulting symbolic invariants are used to create definitions of invariant formulas in Isabelle. For instance, $!(x \doteq true \bar{\wedge} n[1] \doteq C)$ is generalized into $!(x \doteq true \bar{\wedge} n[iInv_1] \doteq C)$. This kind of generalization is done with model constraints, which specify that any parameter index should be not greater than the instance size N , and parameters to instantiate a parameterized rule (formula) should be different. (2) The generalization of concrete causal relations into parameterized causal relations in Isabelle, and will be used in proofs of the existence of causal relations in Isabelle.

Since the first kind of generalization is simple, we focus on the second kind of generalization, which consists of two phases. Firstly, groups of rule parameters such as $[[1],[2],[3]]$ will be generalized into a list of symbolic formulas such as $[iR_1 \doteq iInv_1, iR_1 \doteq iInv_2, (iR_1 \neq iInv_1) \wedge (iR_1 \neq iInv_2)]^3$, which stands for case-splittings by comparing a symbolic rule parameter iR_1 and invariant parameters $iInv_1$ and $iInv_2$. In the second phase, the formula field accompanied with a relation of kind $invHoldRule_3$ is also generalized by some special strategy.

Now let us look at the first phase, starting with some

³ $iR_1 \neq iInv_1$ is the abbreviation of $!(iR_1 \doteq iInv_1)$

definitions. Consider a line of concrete causal relation shown in Table I, there is a group of rule parameters LR , and a group of parameters LI occurring in an invariant formula.

Definition 5 Let LR be a permutation s.t. $|LR| > 0$, which represents a list of actual parameters to instantiate a parameterized rule, let LI be a permutation $|LI| > 0$, which represents a list of actual parameters to instantiate a parameterized invariant, we define:

- 1) symbolic comparison condition generalized from comparing $LR_{[i]}$ and $LI_{[j]}$:

$$\text{symbCmp}(LR, LI, i, j) \equiv \begin{cases} iR_1 \doteq iInv_j & \text{if } LR_{[i]} = LI_{[j]} \\ iR_1 \neq iInv_j & \text{otherwise} \end{cases}$$

- 2) symbolic comparison condition generalized from comparing $LR_{[i]}$ and with all $LI_{[j]}$:

$$\text{symbCaseI}(LR, LI, i) \equiv \begin{cases} \text{symbCmp}(LR, LI, i, j) & \text{if } \exists j. LR_{[i]} = LI_{[j]} \\ \text{forallForm}(|LI|, pf) & \text{otherwise} \end{cases}$$

where $pf(j) = \text{symbCmp}(LR, LI, i, j)$, and $\exists j.P$ is an qualifier meaning that there exists a unique j s.t. property P ;

- 3) symbolic case generalized from comparing LR with LI : $\text{symbCase}(LR, LI) \equiv \text{forallForm}(|LR|, pf)$, where $pf(i) = \text{symbCaseI}(LR, LI, i)$;
- 4) symbolic partition generalized from comparing all $LRS_{[k]}$ with LI , where LRS is a list of permutations with the same length: $\text{partition}(LRS, LI) \equiv \text{existsForm}(|LRS|, pf)$, where $pf(i) = \text{symbCase}(LRS_i, LI)$.

$\text{symbCmp}(LR, LI, i, j)$ defines a symbolic formula generalized from comparing $LR_{[i]}$ and $LI_{[j]}$; $\text{symbCaseI}(LR, LI, i)$ a symbolic formula summarizing the results of comparison between $LR_{[i]}$ and all $LI_{[j]}$ such that $j \leq |LI|$; $\text{symbCase}(LR, LI)$ a symbolic formula representing a subcase generalized from comparing all $LR_{[i]}$ and all $LI_{[j]}$; $\text{partition}(LRS, LI)$ is a disjunction of subcases $\text{symbCase}(LRS_{[i]}, LI)$. Recall the first three lines in Table I, and $LI = [1, 2]$ is the list of parameters occurring in $\text{mutualEx}(1, 2)$; and LR is the actual parameter list to instantiate crit.

- when $LR = [1]$, $\text{symbCmp}(LR, LI, 1, 1) = (iR_1 \doteq iInv_1)$, $\text{symbCase}(LR, LI) = \text{symbCaseI}(LR, LI, 1) = (iR_1 \doteq iInv_1)$ because $LR_{[1]} = LI_{[1]}$.
- when $LR = [2]$, $\text{symbCmp}(LR, LI, 1, 2) = (iR_1 \doteq iInv_2)$, $\text{symbCase}(LR, LI) = \text{symbCaseI}(LR, LI, 2) = (iR_1 \doteq iInv_2)$ because $LR_{[1]} = LI_{[2]}$.
- when $LR = [3]$, $\text{symbCmp}(LR, LI, 1, 1) = (iR_1 \neq iInv_1)$, $\text{symbCmp}(LR, LI, 1, 2) = (iR_1 \neq iInv_2)$, $\text{symbCase}(LR, LI) = \text{symbCaseI}(LR, LI, 1) = (iR_1 \neq iInv_1) \wedge (iR_1 \neq iInv_2)$ because neither $LR_{[1]} = LI_{[1]}$ nor $LR_{[1]} = LI_{[2]}$.

- let $LRS = [[1], [2], [3]]$, $\text{partition}(LRS, LI) = (iR_1 \doteq iInv_1) \vee (iR_1 \doteq iInv_2) \vee ((iR_1 \neq iInv_1) \wedge (iR_1 \neq iInv_2))$

If we see a line in table I as a concrete test case for some concrete causal relation, then $\text{symbCase}(LR, LI)$ is an abstraction predicate to generalize the concrete case. Namely, if we transform $\text{symbCase}(LR, LI)$ by substituting $iInv_i$ with $LI_{[i]}$, and iR_j with $LR_{[j]}$, the result is semantically equivalent to true.

The second phase of generalization of concrete causal relations is to generalize the formula inv' accompanied with a causal relation invHoldRule_3 in a line of table I. An index occurring in f' can either occur in the invariant formula, or in the rule. We need to look it up to determine the transformation.

Definition 6 Let LI and LR are two permutations, $\text{find_first}(L, i)$ returns the least index j s.t. $L_{[j]} = i$ if there exists such an index; otherwise returns an error.

$$\text{lookup}(LI, LR, i) \equiv \begin{cases} iInv_{\text{find_first}(LI, i)} & \text{if } i \in LI \\ iR_{\text{find_first}(LR, i)} & \text{otherwise} \end{cases}$$

$\text{lookup}(LI, LR, i)$ returns the symbolic index transformed from i according to whether i occurs in LI or in LR . The index i will be transformed into $iInv_{\text{find_first}(LI, i)}$ if i occurs in LI , and $iR_{\text{find_first}(LR, i)}$ otherwise. Employing the lookup strategy to transform a concrete index i in inv' to $\text{lookup}(LI, LR, i)$, symbolize2f transforms inv' into a symbolic one which will be needed in a proof command for existence of the invHoldRule_3 relation in Isabelle.

VI. AUTOMATIC GENERATION OF ISABELLE PROOF

A formal model for a protocol case in a theorem prover like Isabelle includes the definitions of constants and rules and invariants, lemmas, and proofs. Readers can refer to [12] for detailed illustration of the formal proof script. In this section, we focus on the generation of a lemma on the existence of causal relation between a parameterized rule and invariant formula based on the aforementioned generalization of lines of concrete causal relations.

An example lemma critVsinv_1 and its proof in Isabelle in the mutualEx protocol, is illustrated as follows:

```

1 lemma critVsinvl:
2   assumes a1:  $\exists iR1. iR1 \leq N \wedge r = \text{crit } iR1$  and
3   a2:  $\exists iInv1\ iInv2. iInv1 \leq N \wedge iInv2 \leq N \wedge iInv1 \neq iInv2$ 
4    $\wedge f = \text{invl } iInv1\ iInv2$ 
5   shows  $\text{invHoldRule } s\ f\ r$  (invariants N)
6   proof -
7     from a1 obtain iR1 where a1:  $iR1 \leq N \wedge r = \text{crit } iR1$ 
8     by blast
9     from a2 obtain iInv1 iInv2 where a2:  $iInv1 \leq N$ 
10     $\wedge iInv2 \leq N \wedge iInv1 \neq iInv2 \wedge f = \text{invl } iInv1\ iInv2$ 
11    by blast
12  have  $iR1 = iInv1 \vee iR1 = iInv2 \vee (iR1 \neq iInv1 \wedge iR1 \neq iInv2)$ 
13  by auto
14  moreover {
15    assume b1:  $iR1 = iInv1$ 
16    have  $\text{invHoldRule3 } s\ f\ r$  (invariants N)
17    proof (cut_tac a1 a2 b1, simp,
18      rule_tac x=!(x=true  $\wedge$  n[iInv2]=C) in exI, auto) qed
19    then have  $\text{invHoldRule } s\ f\ r$  (invariants N) by auto
20  }
21  moreover {
22    assume b1:  $iR1 = iInv2$ 
23    have  $\text{invHoldRule3 } s\ f\ r$  (invariants N)
24    proof (cut_tac a1 a2 b1, simp,
25      rule_tac x=!(x=true  $\wedge$  n[iInv1]=C) in exI, auto) qed
26    then have  $\text{invHoldRule } s\ f\ r$  (invariants N) by auto
27  }
28  moreover {
29    assume b1:  $(iR1 \neq iInv1 \wedge iR1 \neq iInv2)$ 
30    have  $\text{invHoldRule2 } s\ f\ r$ 
31    proof (cut_tac a1 a2 b1, auto) qed
32    then have  $\text{invHoldRule } s\ f\ r$  (invariants N) by auto
33  }
34  ultimately show  $\text{invHoldRule } s\ f\ r$  (invariants N) by blast
35  qed

```

In the above proof, line 2 are assumptions on the parameters of the invariant and rule, which are composed of two parts: (1) assumption a1 specifies that there exists an actual parameter $iR1$ with which r is a rule obtained by instantiating `crit`; (2) assumption a2 specifies that there exists actual parameters $iInv1$ and $iInv2$ with which f is a formula obtained by instantiating `invl`. Line 4 are two typical proof patterns forward-style which fixes local variables such as $iR1$ and new facts such as $a1: iR1 \leq N \wedge r = \text{crit } iR1$. From line 5, the remaining part is a typically readable Isar proof using calculation reasoning such as `moreover` and `ultimately` to do case analysis. Line 5 splits cases of $iR1$ into all possible cases by comparing $iR1$ with $iInv1$ and $iInv2$, which is in fact characterized by `partition([1], [2], [3], [1, 2])`. Lines 6-14 proves these cases one by one: Lines 6-8 proves the case where $iR1 = iInv1$, line 7 first proves that the causal relation invHoldRule_3 holds by supplying a symbolic formula, which is transformed from $\text{invOnXC}(2)$ by calling `symbolize2f` with `lookUp` strategy. From the conclusion at line 7, line 8 furthermore proves the causal relation invHoldRule holds; Lines 9-11 proves the case where $iR1 = iInv2$, proof of which is similar to that of case 1; Lines 12-14 the case where neither $iR1 = iInv1$ nor $iR1 = iInv2$. Each proof of a subcase is done in a block `moreover b1: asml proof1`, the `ultimately` proof command in line 15 concludes by summing up all the subcases.

Due to length limitation, we illustrate the algorithm for generating a key part of the proof of the lemma `critVsinvl`: the generation of a subproof (e.g., lines 7-8) according to a symbolic relation tag of invHoldRule_{1-3} , which is shown in Algorithm 3. Input relTag is the result of the generalization step, which is discussed in Section V. In the body of function `rel2proof`, `sprintf` writes a formatted data to string and returns it. In line 10, `getFormField(relTag)` returns the field of formula f' if $\text{relTag} = \text{invHoldRule}_3(f')$. `rel2proof` transforms

Algorithm 3: Generating a kind of proof which is according with a relation tag of $\text{invHoldRule}_{1-3} : \text{rel2proof}$

Input: A symbolic causal relation item relTag

Output: An Isabelle proof: proof

```

1 if  $\text{relTag} = \text{invHoldRule}_1$  then
2    $\text{proof} \leftarrow \text{sprintf}$ 
3   "have  $\text{invHoldRule1 } f\ r$  (invariants N)
4   by (cut_tac a1 a2 b1, simp, auto)
5   then have  $\text{invHoldRule } f\ r$  (invariants N) by blast" ;
6 else if  $\text{relTag} = \text{invHoldRule}_2$  then
7    $\text{proof} \leftarrow \text{sprintf}$ 
8   "have  $\text{invHoldRule2 } f\ r$  (invariants N) by (cut_tac a1 a2
9   b1, simp, auto)
10  then have  $\text{invHoldRule } f\ r$  (invariants N) by blast" ;
11 else
12    $f' \leftarrow \text{getFormField}(\text{relTag})$ ;
13    $\text{proof} \leftarrow \text{sprintf}$ 
14   "have  $\text{invHoldRule3 } f\ r$  (invariants N)
15   proof (cut_tac a1 a2 b1, simp, rule_tac x=%s in
16   exI, auto) qed
17   then have  $\text{invHoldRule } f\ r$  (invariants N) by blast"
18   (symbf2Isabelle f)";
19 return  $\text{proof}$ 

```

a symbolic relation tag into a paragraph of proof, as shown in lines 7-8, 10-11, or 13-14. If the tag is among invHoldRule_{1-2} , the transformation is rather straight-forward, else the form f' is assigned by the formula `getFormField(relTag)`, and provided to tell Isabelle the formula which is used to construct the invHoldRule_3 relation.

VII. EXPERIMENTS

We implement our tool in Ocaml. Experiments are done with typical bus-snoopy benchmarks such as MESI and MOESI, as well as directory-based benchmarks such as German and FLASH. The detailed codes and experiment data can be found in [12]. Each experiment data includes the `paraVerifier` instance, invariant sets, Isabelle proof scripts. Experiment results are summarized in Table II.

Among all the work in the field of parameterized verification, only four of them have verified FLASH. The first full verification of safety properties of FLASH is done in [10]. Park and Dill proved the safety properties of FLASH using PVS. The CMP method, which adopts parameter abstraction and guard strengthening, is applied in [6] for verifying safety properties of FLASH. McMillan applied compositional model checking [11] and used Candence SMV to the verification of both safety and liveness properties of FLASH. Sylvain et.al have applied Cubeic to the verification FLASH [9], [13], which is theoretically based on an SMT model checking to the verification of array-based system. In the former three methods [10], [6], [11], auxiliary invariants are provided manually depending on verifier's deep insight in the FLASH protocol itself, while in Cubeic, auxiliary invariants are found automatically. In Cubeic, auxiliary invariants are searched backward by a heuristics-guided algorithm with the help of an

TABLE II
VERIFICATION RESULTS ON BENCHMARKS.

Protocols	#rules	#invariants	time (seconds)	Memory (MB)
mutualEx	4	5	3.25	7.3
MESI	4	3	2.47	11.5
MOESI	5	3	2.49	23.2
Germanish [9]	6	3	2.9	7.8
German [6]	13	52	38.67	14
FLASH_nodata	60	152	280	26
FLASH_data	62	162	510	26

oracle (a reference instance of the protocol), but these auxiliary invariants are in concrete form, and are not generalized to the parameterized form. Thus there is no parameterized proof derived for parameterized verification of FLASH.

The invariants-searching algorithm used in our work differs from that in Cubeic [9], [13] in that the heuristics in our work are based on the construction of causal relation which is uniquely proposed in our work. Thus the auxiliary invariants in our work are different from those found in [9], [13]. Moreover, we generalize these concrete invariants and causal relations into a parameterized proof, and generate a parameterized proof in Isabelle. The found invariants have abundant semantics reflecting the deep insight of the FLASH protocol design, and the readable Isabelle proof script formally proves these invariants. In this way, we prove the protocol with the highest assurance. To the best of knowledge, this work for the first time automatically generates a proof of safety properties of full version of FLASH in a theorem prover without auxiliary invariants manually provided by people.

VIII. CONCLUSION

The originality of paraVerifier lies in the following aspects: (1) instead of directly proving the invariants of a protocol by induction, we propose a general proof method based on the consistency lemma to decompose the proof goal into a number of small ones; (2) instead of proving the decomposed subgoals by hand, we automatically generate proofs for them based on the information of causal relation computed in a small protocol instance.

As we demonstrate in this work, combining theorem proving with automatic proof generation is promising in the field of formal verification of industrial protocols. Theorem proving can guarantee the rigorousness of the verification results, while automatic proof generation can release the burden of human interaction.

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