Software Model Checking: The VeriSoft Approach

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Abstract

Verification by state-space exploration, also often referred to as *model checking*, is an effective method for analyzing the correctness of concurrent reactive systems (for instance, communication protocols). Unfortunately, traditional model checking is restricted to the verification of properties of models, i.e., *abstractions*, of concurrent systems.

We discuss in this paper how model checking can be extended to analyze arbitrary *software*, such as implementations of communication protocols written in programming languages like C or C++. We then introduce a search technique that is suitable for exploring the state spaces of such systems. This algorithm has been implemented in *VeriSoft*, a tool for systematically exploring the state spaces of systems composed of several concurrent processes executing arbitrary code.

During the past five years, VeriSoft has been applied successfully for analyzing several software products developed in Lucent Technologies, and has also been licensed to hundreds of users in industry and academia. We discuss applications, strengths and limitations of VeriSoft, and compare it to other approaches to software model checking, analysis and testing.

1 Introduction

Concurrent systems are systems composed of elements that can operate concurrently and communicate with each other. Each component can be viewed as a *reactive system*, i.e., a system that continuously interacts with its environment. Concurrent reactive systems are notably hard to design and test because their components may interact in many unexpected ways. Traditional testing techniques are of limited help since test coverage is bound to be only a minute fraction of the possible behaviors of the system. Furthermore, scenarios leading to errors are often extremely difficult to reproduce.

An effective approach for analyzing the correctness of a concurrent reactive system consists of *systematically exploring its state space*. The state space of a concurrent system is a directed graph that represents the combined behavior of all concurrent components in the system. Such a state space can be computed automatically from an abstract description of the concurrent system specified in a (essentially finite-state) *modeling language*. Many properties of a model of a system

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can be checked by exploring its state space: deadlocks, dead code, violations of user-specified assertions, etc. Moreover, the range of properties that state-space exploration techniques can verify has been substantially broadened during the last decade thanks to the development of model-checking methods for various temporal logics (e.g., [CES86, LP85, QS81, VW86]). In what follows, we will use the term "model checking" in a broad sense, to denote any automatic state-space exploration technique that can be used for verification purposes.[1]

Examples of tools that follow the above paradigm are CAESAR [FGM$^+$92], COSPAN [HK90], CWB [CPS93], MURPHI [DDHY92], SMV [McM93] and SPIN [Hol91], among others. These tools differ by the modeling languages they use for representing systems and properties. But all of them are based on state-space exploration algorithms, in one form or another, for performing the verification itself.

The effectiveness of model checking for analyzing the correctness of concurrent reactive systems is becoming increasingly well-established. A large variety of complex reactive systems, ranging from circuit designs to communication protocols, have been modeled and then analyzed using state-space exploration techniques. In many cases, these analyzes were able to reveal quite subtle design errors (e.g., [Rud92, CGH$^+$93, BG96]). Once the model of a new software application has been thoroughly analyzed, it can also be used as the core of the implementation of the application (e.g., [HP89, FHS95]).

It is worth emphasizing that the practical interest of systematic state-space exploration (and of "verification" in general) is mainly to find errors that would be hard to detect and reproduce otherwise, and not necessarily to prove the absence of errors. While mathematically proving that a model of a system conforms to a specific set of properties does increase the confidence that the actual system is "correct", it does not provide a proof of this fact. Therefore, although model checking is a verification framework, it is closer to *testing* in practice since any verification process is inherently incomplete: only some abstract models or system configurations can be checked against some properties in some environment, and verification results can also be approximate when an exact answer is too expensive to compute.

In this paper, we discuss how model checking can be extended to deal directly with "actual" descriptions of concurrent systems, such as implementations of communication protocols written in general-purpose programming languages like C or C++. We show that existing search techniques for state-space exploration are fundamentally limited to the analysis of systems for which each state of the system can be readily represented by a unique identifier. We then introduce an efficient search technique that does not rely on this assumption. This search algorithm can therefore be applied to systems composed of several concurrent processes executing arbitrary code written in full-fledged programming languages.

[1] Traditionally, the term "model checking" means "to check whether a system is a model of a temporal logic formula", in the classic logical sense. This definition does not imply that a "model", i.e., an abstraction, of a system is checked.

2 Concurrent Systems and Dynamic Semantics

We consider a concurrent system composed of a finite set $\mathcal{P}$ of *processes* and a finite set $\mathcal{O}$ of *communication objects*. Each process $P \in \mathcal{P}$ executes a sequence of *operations* described in a sequential program written in a full-fledged programming language such as C or C++. Such sequential programs are *deterministic*: every execution of the program on the same input data performs the same sequence of operations. We assume that processes communicate with each other by performing atomic operations on communication objects. A communication object $O \in \mathcal{O}$ is defined by a pair (V, OP), where V is the set of all possible values for the object (its domain), and OP is the set of *operations* that can be performed on the object. Examples of communication objects are shared variables, semaphores, and FIFO buffers. Since we assume operations on communication objects are atomic, at most one operation can be performed on a given communication object at any time. Operations on communication objects are called *visible operations*, while other operations are by default called *invisible*. The execution of an operation is said to be *blocking* if it cannot currently be completed; for instance, waiting for the reception of a message blocks until a message is received. We assume that only executions of visible operations may be blocking.

At any time, the concurrent system is said to be in a *state*. The system is said to be in a *global state* when the next operation to be executed by every process in the system is a visible operation. Every process in the system is always expected to eventually attempt executing a visible operation. (If a process does not attempt to execute a visible operation within a given amount of time, an error, called divergence, is reported at run-time.) This assumption implies that initially, after the creation of all the processes of the system, the system can reach a first and unique global state s_0, called the *initial global state* of the system. A *process transition*, or *transition* for short, is defined as one visible operation followed by a *finite* sequence of invisible operations performed by a single process and ending just before a visible operation. Let $\mathcal{T}$ denote the set of all transitions of the system.

A transition is said to be *disabled* in a global state s when the execution of its visible operation is blocking in s. Otherwise, the transition is said to be *enabled* in s. A transition t that is enabled in a global state s can be *executed* from s. Since the number of invisible operations in a transition is finite, the execution of an enabled transition always terminates. When the execution of t from s is completed, the system reaches a global state s', called the *successor* of s by t.[2] We write $s \xrightarrow{t} s'$ to mean that the execution of the transition t leads from the global state s to the global state s', while $s \overset{w}{\Rightarrow} s'$ means that the execution of the finite sequence w of transitions leads from s to s'.

We now define a formal semantics for the concurrent systems that satisfy our assumptions. A concurrent system as defined here is a closed system: from its initial global state, it can evolve and change its state by executing enabled transitions. Therefore, a natural way to describe the possible *behaviors* of such a system is to consider its set of reachable global states and the transitions that are possible between these.

Formally, the joint *global* behavior of all processes P_i in a concurrent system can be represented

[2] Operations on objects (and hence transitions) are deterministic: the execution of a transition t in a state s leads to a *unique* successor state.

```
/* phil.c : dining philosophers (version without loops) */

#include <stdio.h>
#include <sys/types.h>
#include <sys/ipc.h>
#include <sys/sem.h>

#define N 2

philosopher(i,semid)
     int i, semid;
{
  printf("philosopher %d thinks\n",i);
  semwait(semid,i,1);          /* take left fork */
  semwait(semid,(i+1)%N,1);    /* take right fork */
  printf("philosopher %d eats\n",i);
  semsignal(semid,i,1);       /* release left fork */
  semsignal(semid,(i+1)%N,1); /* release right fork */
  exit(0);
}

main()
{
  int semid, i, pid;

  semid = semget(IPC_PRIVATE,N,0600);

  for(i=0;i<N;i++)
    semsetval(semid,i,1);
  for(i=0;i<(N-1);i++) {
    if((pid=fork()) == 0)
      philosopher(i,semid);
    };
  philosopher(i,semid);
}
```

Figure 1: Example of concurrent C program simulating dining philosophers

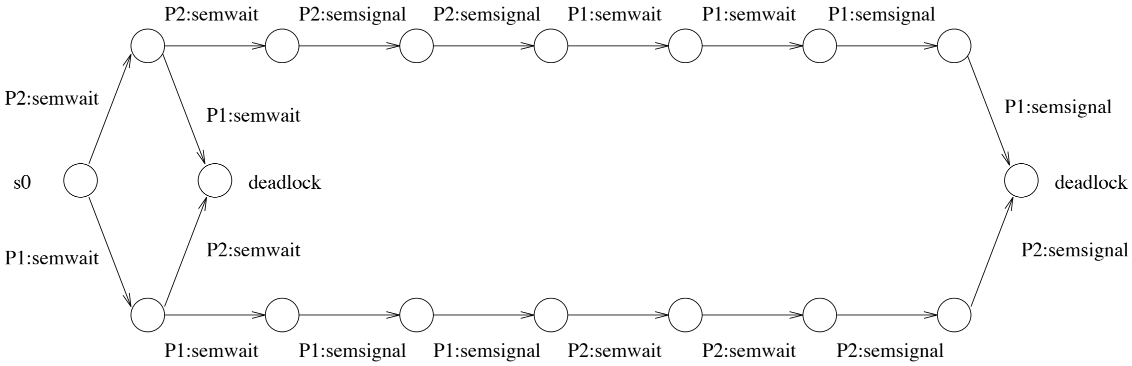


Figure 2: Global state space for the two-dining-philosophers system

by a transition system $A_G = (S, \Delta, s_0)$ such that

- S is the set of global states of the system,
- $\Delta \subseteq S \times S$ is the *transition relation* defined as follows:

$$(s, s') \in \Delta \text{ iff } \exists t \in \mathcal{T} : s \xrightarrow{t} s',$$

- s_0 is the initial global state of the system.

An element of Δ corresponds to the execution of a single transition $t \in \mathcal{T}$ of the system. The elements of Δ will be referred to as *global transitions.* As usual, we restrict A_G to its global states and transitions that are reachable from s_0, since the other global states and transitions play no role in the behavior of the system. In what follows, a "state in A_G" denotes a state that is reachable from s_0. By definition, states in A_G are global. We call A_G the *global state space* of the system.

Example 1 Consider the concurrent C program shown in Figure 1. This program represents a concurrent system composed of two processes that communicate using UNIX semaphores. The program describes the behavior of these processes as well as the initialization of the system. This example is inspired by the well-known dining-philosophers problem, with two philosophers. The two processes communicate by executing the (visible) operations *semwait* and *semsignal* on two semaphores that are identified by the integers 0 and 1 respectively. The operations *semwait* and *semsignal* take 3 arguments: the first argument is an identifier for an array of semaphores, the second is the index of a particular semaphore in that array, and the third argument is a value by which the counter associated with the semaphore identified by the first two arguments must be decremented (in the case of *semwait*) or incremented (in the case of *semsignal*). The value of both semaphores is initialized to 1 with the operation *semsetval.* By implementing these operations using actual UNIX semaphores (the exact UNIX system calls to do this are similar), the program above can be compiled and run on any UNIX machine. The state space A_G of this system is shown in Figure 2, where the two processes are denoted by $P1$ and $P2$, and global transitions are labeled with the visible operation of the corresponding process transition. The operation *exit* is a visible operation whose execution is always blocking. Since all the processes are deterministic, nondeterminism in A_G is caused only by concurrency. ■

Since we consider here closed concurrent systems, the environment of one process is formed by the other processes in the system. This implies that, in the case of a single "open" reactive system, the environment in which this system operates has to be represented, possibly using other processes. In practice, a complete representation of such an environment may not be available, or may be very complex. It is then convenient to use a simplified representation of the environment to simulate its behavior. For this purpose, we introduce a special operation "VS_toss" to express a valuable feature of modeling languages, not found in programming languages: *nondeterminism.* This operation takes as argument a positive integer n, and returns an integer in $[0, n]$. The operation is visible and nondeterministic: the execution of a transition starting with VS_toss(n) may yield up to $n + 1$ different successor states, corresponding to different values returned by VS_toss.

Which properties of a concurrent system is it possible to check by examining its state space A_G as defined above? Here, we focus mainly on two verification problems (other properties will be discussed later in Section 5): the detection of *deadlocks* and *assertion violations.* Deadlocks are states where the execution of the next operation of every process in the system is blocking. Deadlocks are a notorious problem in concurrent systems, and are extremely difficult to detect through conventional testing. Assertions can be specified by the user with the special operation "VS_assert". This operation can be inserted in the code of any process, and is considered visible. It takes as its argument a boolean expression that can test and compare the value of variables and data structures local to the process. When "VS_assert(expression)" is executed, the expression is evaluated. If the expression evaluates to false, the assertion is said to be *violated.* Many undesirable system properties, such as unexpected message receptions, buffer overflows and application-specific error conditions, can easily be expressed as assertion violations.

The following theorem states that deadlocks and assertion violations can be detected by exploring only the global states of a concurrent system.

Theorem 1 *Consider a concurrent system as defined above, and let A_G denote its state space. Then, all the deadlocks that are reachable after the initialization of the system are global states, and are therefore in A_G. Moreover, if there exists a state reachable after the initialization of the system where an assertion is violated, then there exists a global state in A_G where the same assertion is violated.*

Proof: See Appendix. ■

This theorem justifies our choice for the "dynamic" semantics described in this section.

In the next section, we discuss how to build a representation of the state space of a concurrent system as defined above. We briefly review standard state-space exploration techniques, and show why they are not appropriate for exploring state spaces of concurrent systems whose processes are described by arbitrary programs.

```
1   Initialize: $Unexplored$ is empty; $Explored$ is empty;
2               add $s_0$ to $Unexplored$;
3   Loop: while $Unexplored \neq \emptyset$ do {
4               take $s$ out of $Unexplored$;
5               if $s$ is NOT already in $Explored$ then {
6                  enter $s$ in $Explored$;
7                  $T = enabled(s)$;
8                  for all $t$ in $T$ do {
9                       $s'$ = succ($s$) after $t$;
10                      add $s'$ to $Unexplored$;
11                      }
12                 }
13              }
```

Figure 3: Algorithm 1 – classical search

3 Existing State-Space Exploration Techniques

In the case of *models* of concurrent systems, a state space A_G is usually computed by performing a search of all the states that are reachable from the initial state s_0 of the model of the system. An algorithm for performing such a search is shown in Figure 3. This algorithm recursively explores all successor states of all states encountered during the search, starting from the initial state, by executing all enabled transitions in each state (lines 7–8). The main data structures used are a set *Unexplored* to store the states whose successors still have to be explored, and a set *Explored* (often implemented as a hash-table) to store all the states that have already been visited during the search. The set of all transitions enabled in a state s is denoted by $enabled(s)$. The state reached from a state s after the execution of a transition t is denoted "succ(s) after t". It is easy to prove that, if A_G is finite, all the states of A_G are visited during the search performed by the algorithm of Figure 3 [AHU74]. The order in which the search is performed (depth-first, breadth-first, etc.) depends on how the operations "add" and "take" are implemented.

It is important to note that the algorithm of Figure 3 assumes that each state s can be represented by a *unique identifier*, that can be stored in the data structures *Unexplored* and *Explored* during the search. Although other search algorithms for modeling languages, such as symbolic verification methods [BCM$^+$90, CGL92, McM93], may use other types of data structures (e.g., Binary Decision Diagrams [Bry92]) for representing state spaces, they all rely on the assumption that each state of the system has a unique representation (typically a string of bits) that is easy to compute and manipulate.

When dealing with processes described by arbitrary programs written in full-fledged programming languages, this assumption is not valid anymore. Indeed, the state of each process is determined by the values of all the memory locations that can be accessed by the process and influence its behavior (including activation records associated with procedure calls). This information is typically far too large and complex to be efficiently and unambiguously encoded by a string of bits, which could then be saved in memory at each step of the state-space exploration.

However, nothing prevents us from systematically searching the state space of a concurrent system without storing any intermediate states in memory, by successively enumerating and exploring all possible sequences of transitions in the state space. Let us call such a search a *state-less search*. Of course, if the state space A_G contains cycles, a state-less search through it will not terminate, even if A_G is finite (unless an upper bound on the number of states of A_G is known). Even state-less searches of "small" finite acyclic state spaces (e.g., composed of only a few thousand states) may not terminate in a reasonable amount of time. To illustrate this phenomenon, let us consider the dining-philosophers example again. (The state space of this system does not contain any cycles.) The number of transitions explored by a classical search (Algorithm 1) and by a state-less search are compared in Figure 5, for various numbers N of processes. The run-time of both algorithms is proportional to the number of explored transitions. One clearly sees that the state-less search is much slower than the classical one. In the case of four processes, the state-less search explores 386,816 transitions, while they are only 708 transitions in A_G. While every transition of A_G is executed exactly once during a classical search, every transition of A_G is executed on average about 546 times during a state-less search! This tremendous difference is due to the numerous re-explorations of unstored parts of the state space during the state-less search.

4 An Efficient State-Less Search Algorithm

The state-less search technique can be viewed as a particular case of *state-space caching* [Hol85, JJ91, GHP95], a memory management technique for storing the states encountered during a classical search performed in depth-first order. State-space caching consists of storing all the states of the current explored path plus as many other states as possible given the remaining amount of available memory. It thus creates a restricted *cache* of selected states that have already been visited. This method never tries to store more states than possible in the cache. A state-less search corresponds to the extreme case where the cache does not contain any state at all.

State-space caching suffers the same drawback as the state-less search: multiple redundant explorations of large unstored parts of the state space yield an unacceptable blow-up of the run-time. Indeed, almost all states in the state spaces of concurrent systems are typically reached several times during the search. There are two causes for this:

1. From the initial state, the exploration of any interleaving of a single finite partial ordering of transitions of the system always leads to the same state. This state will thus be visited several times because of all these interleavings.

2. From the initial state, explorations of different finite partial orderings of transitions may lead to the same state.

In [GHP95], it is shown that most of the effects of the first cause given above can be avoided when using a search algorithm based on the notion of *sleep sets* [God90, GW93]. Such an algorithm dynamically prunes the state space of a concurrent system without incurring the risk of any incompleteness in the verification results. Empirical results [GHP95, God96] show that, in many cases, most of the states are visited *only once* during a state-space exploration performed with this search technique. This makes it possible not to store most of the states previously visited during the search without incurring much redundant exploration of parts of the state space.

Sleep sets belong to a broader family of algorithms, referred to as *partial-order methods* [God96], that were developed to tackle the "state explosion" phenomenon that limits the efficiency and applicability of verification by state-space exploration. In [God96], it is shown that sleep sets can be combined with another pruning technique based on the notion of *persistent sets*. Using both techniques simultaneously preserves the beneficial properties of sleep sets outlined in the previous paragraph while substantially reducing the number of states and transitions that have to be visited.

In this section, we present a new state-space exploration algorithm that combines a state-less search with the persistent-set and sleep-set techniques. Before turning to the presentation of this algorithm, we briefly recall some basic principles of partial-order methods.

The basic idea behind partial-order methods that enables them to check properties of A_G without constructing the whole of A_G is the following: A_G contains many paths that correspond simply to different execution orders of the same process transitions. If these transitions are "independent", for instance because they are executed by noninteracting processes, then changing their order will not modify their combined effect.

This notion of independency between transitions and its complementary notion, the notion of dependency, can be formalized by the following definition (adapted from [KP92]).

Definition 1 Let $\mathcal{T}$ be the set of system transitions and $D \subseteq \mathcal{T} \times \mathcal{T}$ be a binary, reflexive, and symmetric relation. The relation D is a *valid dependency relation* for the system iff for all $t_1, t_2 \in \mathcal{T}$, $(t_1, t_2) \notin D$ (t_1 and t_2 are *independent*) implies that the two following properties hold for all global states s in the global state space A_G of the system:

1. if t_1 is enabled in s and $s \overset{t_1}{\rightarrow} s'$, then t_2 is enabled in s iff t_2 is enabled in s' (independent transitions can neither disable nor enable each other); and
2. if t_1 and t_2 are enabled in s, then there is a unique state s' such that $s \overset{t_1 t_2}{\Rightarrow} s'$ and $s \overset{t_2 t_1}{\Rightarrow} s'$ (commutativity of enabled independent transitions).

■

This definition characterizes the properties of possible "valid" dependency relations for the transitions of a given system. In practice, it is possible to give easily checkable syntactic conditions that are sufficient for transitions to be independent. In a concurrent system as defined in Section 2, dependency can arise between transitions of different processes that perform (visible) operations on the same communication objects. For instance, two wait operations on a binary semaphore are dependent when they are enabled, while two signal operations on a same non-binary semaphore are independent. Carefully defining dependencies between operations on communication objects is by no means a trivial task. We refer the reader to [God96] for a detailed presentation of that topic.

All partial-order algorithms follow the same basic pattern: they operate as classical state-space searches except that, at each state s reached during the search, they compute a subset T of the set of transitions enabled at s, and explore only the transitions in T, the other enabled transitions are not explored. Such a search is called a *selective search*. It is easy to see that a selective search through A_G only reaches a subset (not necessarily proper) of the states and transitions of A_G.

Two main techniques for computing such sets T have been proposed in the literature: the persistent-set and sleep-set techniques. The first technique actually corresponds to a whole family of algorithms [Ove81, Val91, GP93, GW93, Pel93]. In [God96], it is shown that all these algorithms compute "persistent sets". Intuitively, a subset T of the set of transitions enabled in a state s of A_G is called *persistent in s* if all transitions not in T that are enabled in s, or in a state reachable from s through transitions not in T, are independent with all transitions in T. In other words, whatever one does from s, while remaining outside of T, does not interact with or affect T. Formally, we have the following [GP93].

Definition 2 A set T of transitions enabled in a state s is *persistent in s* iff, for all nonempty sequences of transitions

$$s = s_1 \xrightarrow{t_1} s_2 \xrightarrow{t_2} s_3 \ldots \xrightarrow{t_{n-1}} s_n \xrightarrow{t_n} s_{n+1}$$

from s in A_G and including only transitions $t_i \notin T$, $1 \leq i \leq n$, t_n is independent with all transitions in T. ■

Note that the set of all enabled transitions in a state s is trivially persistent since nothing is reachable from s by transitions that are not in this set. It is beyond the scope of this paper to present algorithms for computing persistent sets. In a nutshell, these algorithms infer persistent sets from the static structure of the system being verified (such as "process x can perform operation y on communication object z"). They differ by the type of information about the system that they use. The aim of these algorithms is to obtain the smallest possible nonempty persistent sets. See [God96] for several such algorithms and a comparison of their complexity.

The second technique for computing the set of transitions T to consider in a selective search is the sleep set technique [God90, GW93]. This technique does not exploit information about the static structure of the system, but rather about the past of the search. Used in conjunction with a persistent set algorithm, sleep sets can further reduce the number of explored states and transitions.

An algorithm that combines persistent sets and sleep sets with a state-less search is shown in Figure 4. This algorithm performs a selective depth-first search (DFS) in the state space of a concurrent system. The data structure *Stack* contains the sequence of transitions that leads from the initial global state s_0 to the current global state being explored. A set denoted by *Sleep* is associated with each global state reached during the search, i.e., with each call to the procedure DFS. The sleep set associated with a global state s is a set of transitions that are *enabled* in s but *will not be explored* from s. The sleep set associated with the initial global state s_0 is the empty set. Each time a new global state s is encountered during the search, a call to DFS is executed. The sleep set that is associated with s is passed as argument. In line 6, a new set of transitions is selected to be explored from s. Persistent_Set() returns a persistent set in the current global state s that is nonempty if there exist transitions enabled in s. Lines 11 and 14 describe how to

```
1    Initialize: $Stack$ is empty;
2    Search() {
3       DFS($\emptyset$);
4       }
5    DFS(set: $Sleep$) {
6       $T$ = Persistent_Set()$\setminus Sleep$;
7       while $T \neq \emptyset$ do {
8           take $t$ out of $T$;
9           push ($t$) onto $Stack$;
10          Execute($t$);
11          DFS($\{t' \in Sleep \mid t'$ and $t$ are independent$\}$);
12          pop $t$ from $Stack$;
13          Undo($t$);
14          $Sleep = Sleep \cup \{t\}$;
15          };
16      }
```

Figure 4: Algorithm 2 – state-less depth-first search using persistent sets and sleep sets

compute the sleep sets associated with the successor global states of s from the value of its sleep set *Sleep*. In line 10, a transition t is executed from s. The procedure Execute(t) returns after a new global state has been reached by the concurrent system. Then all the transitions of *Sleep* that are independent with t are passed into the sleep set associated with that new global state (line 11). Once the search from that new state (and hence the corresponding call to DFS) is completed, the exploration of the other transitions selected to be explored from s may proceed. The concurrent system is then brought back to the global state s in line 13. (This can be done by reinitializing the system and reexecuting the sequence of transitions in *Stack*, for instance.) Next, transition t, i.e., the last transition explored from s, is added to *Sleep* in line 14.

The correctness of Algorithm 2 is established by the following theorem.

Theorem 2 *Consider a concurrent system as defined in Section 2, and let A_G denote its state space. Assume A_G is finite and acyclic. Then, all the deadlocks in A_G are visited by Algorithm 2. Moreover, if there exists a global state in A_G where an assertion is violated, then there exists a global state visited by Algorithm 2 where the same assertion is violated.*

Proof: See Appendix. ■

In other words, deadlocks and assertion violations can be detected using Algorithm 2. As discussed in the previous section, the termination of Algorithm 2 is guaranteed only when the state space A_G is finite and does not contain any cycles. In practice, Algorithm 2 can be used to efficiently explore the state space of any concurrent system, whether its state space is acyclic or not.

Finally note that Algorithm 2 is different from the algorithms combining persistent sets and

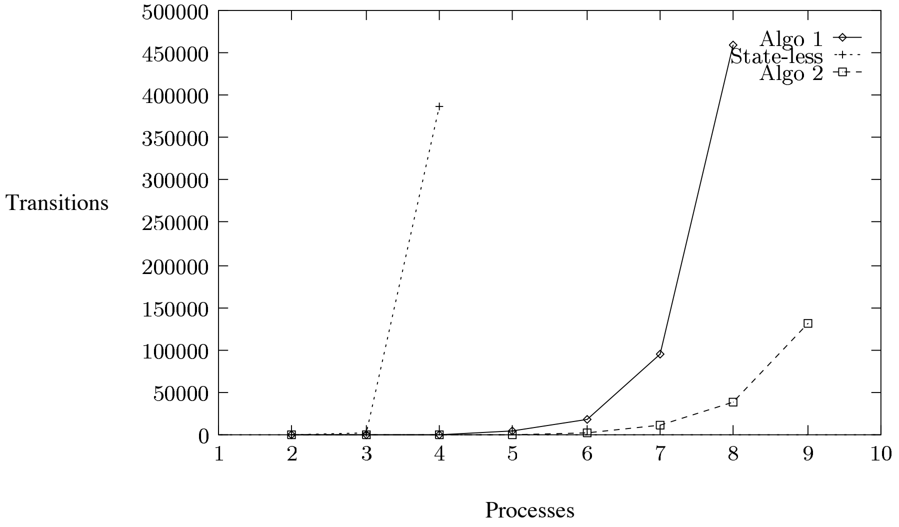


Figure 5: Comparison of performances for the dining-philosophers system

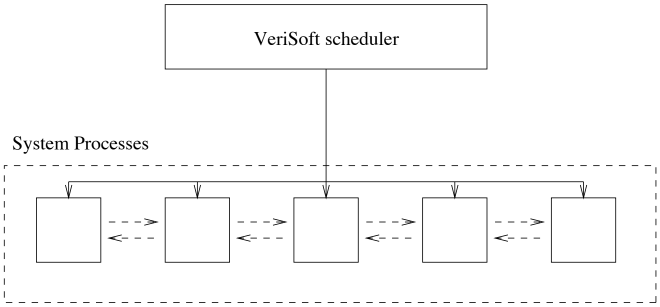


Figure 6: Overall architecture of VeriSoft (in automatic state-space exploration mode)

sleep sets that appeared in [God96]. Indeed, with a state-less search, different sleep sets associated with the same global state (corresponding to different visits of that state via different paths from s_0) cannot interfere with each other during the search. Moreover, cycles cannot be detected in the context of a state-less search, which makes the use of the provisos discussed in [God96] impossible.

Results of experiments with Algorithm 2 for the dining-philosophers example are presented in Figure 5. Thanks to the use of persistent sets and sleep sets, the run-time explosion of the state-less search is now avoided. Moreover, they yield a significant reduction in the number of transitions that need be explored. Although Algorithm 2 does not store any state in memory, it explores fewer transitions than Algorithm 1!

5 VeriSoft

VeriSoft is a tool for systematically exploring the state space of a concurrent system as defined in Section 2. Systematic state-space exploration is performed by controlling and observing the execution of all the visible operations of the concurrent processes of the system. Every process of the concurrent system to be analyzed is mapped to a UNIX process. The execution of the system processes is controlled by an external process, called the *scheduler* (see Figure 6). This process observes the visible operations performed by processes inside the system, and can suspend their execution. By resuming the execution of (the next visible operation of) one selected system process in a global state, the scheduler can explore one transition in the state space A_G of the concurrent system. The scheduler also contains an implementation of a state-less search using persistent sets and sleep sets similar to Algorithm 2. In order to prevent the state-less search from looping forever in cycles of the state space being explored, the depth of the search is limited. When this maximum depth is reached, the scheduler reinitializes the system in order to explore alternative paths in the state space.

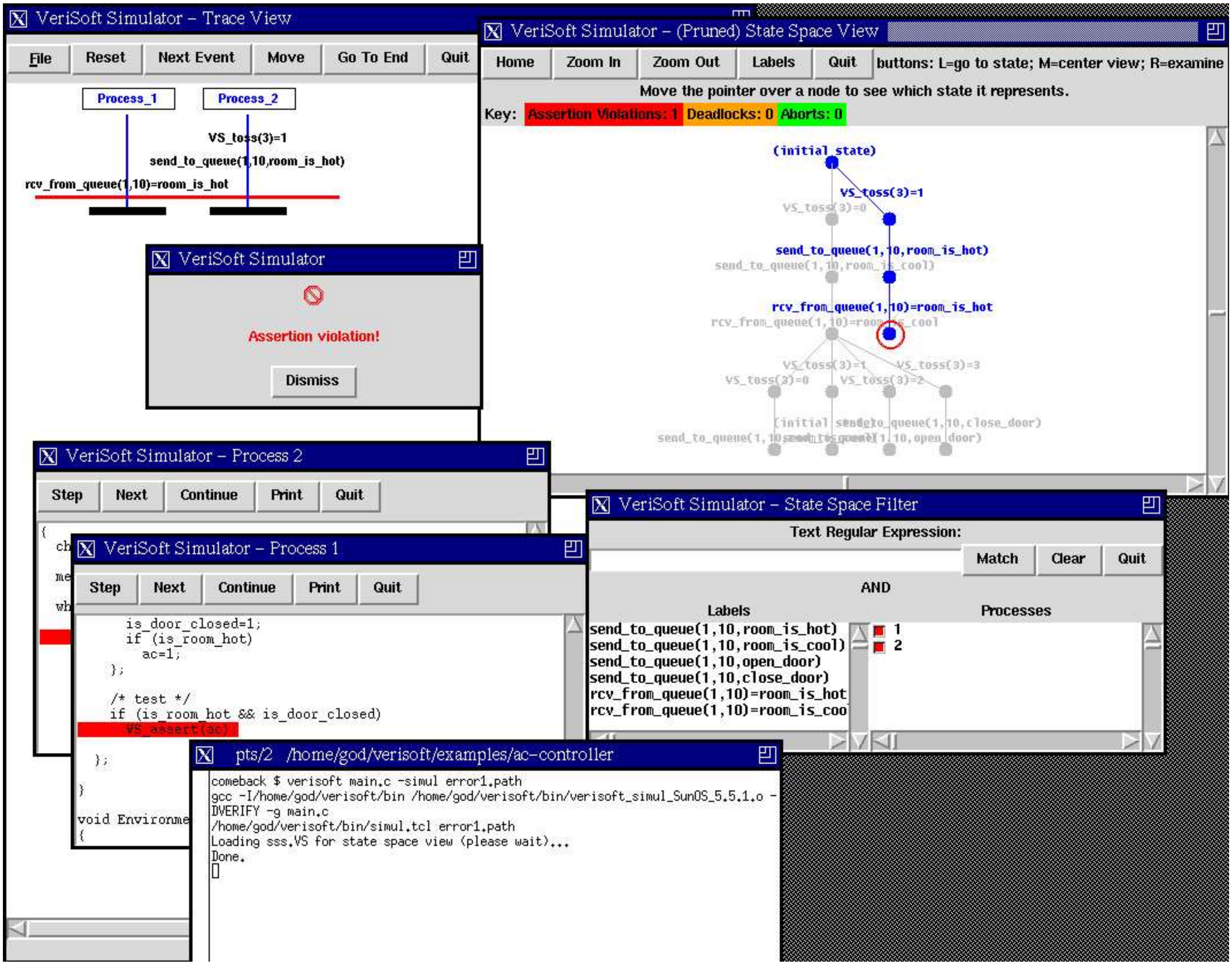
When a deadlock or an assertion violation is detected, the search is stopped, and a scenario formed by all the transitions currently stored in *Stack* (see Figure 4) is exhibited to the user. An interactive graphical simulator/debugger is also available for replaying scenarios and following their executions at the instruction or procedure/function level (see Figure 7). Values of variables of each process can be examined interactively. Using the VeriSoft simulator, the user can also explore any path in the state space of the system with the same set of debugging tools.

It is thus assumed that there are exactly two sources of nondeterminism in the concurrent systems considered here: concurrency and calls to the special visible operation VS_toss used to model nondeterminism as described in Section 2. When this assumption is satisfied, the VeriSoft scheduler has complete control over nondeterminism. It can thus reproduce any scenario leading to an error found during a state-space search and can also guarantee, from a given initial state, complete coverage of the state space up to some depth.

In addition to deadlocks and assertion violations, VeriSoft also checks for *divergences* and *livelocks*. A "divergence" occurs when a process does not attempt to execute any visible operation within a given (user-specified) amount of time. Divergences may be caused by segmentation faults, non-terminating loops, etc. A "livelock" occurs when the execution of the next visible operation of some process is blocking during a sequence of more than a given (user-specified) number of successive states in the state space. Note that these definitions of divergence and livelock differ from the standard definitions for these notions, which correspond to *liveness* properties, i.e., properties that can only be violated by *infinite* sequences of operations or transitions [Lam77, MP92]. In contrast, our notions of divergence and livelock can be violated by *finite* sequences of operations or transitions, and therefore are actually *safety* properties. Indeed, a state-less search cannot detect cycles, and is thus restricted to the verification of safety properties.

6 Applications

During the last five years, VeriSoft has been applied successfully for analyzing several software products developed in Lucent Technologies. Some of these projects have been documented in published papers. For instance, [GHJ98] describes the analysis of the "Heart-Beat Monitor" of a 4ESS switch, a critical component of a telephone switch, while [CGP02] reports on the analysis of several


VeriSoft Simulator – Trace View
File
Reset
Next Event
Move
Go To End
Quit
Process_1
Process_2
VS_toss(3)=1
send_to_queue(1,10,room_is_hot)
rcv_from_queue(1,10)=room_is_hot
VeriSoft Simulator – (Pruned) State Space View
Home
Zoom In
Zoom Out
Labels
Quit
buttons: L=go to state; M=center view; R=examine
Key:
Assertion Violations: 1
Deadlocks: 0
Aborts: 0
(initial state)
VS_toss(3)=1
VS_toss(3)=0
send_to_queue(1,10,room_is_hot)
send_to_queue(1,10,room_is_cool)
rcv_from_queue(1,10)=room_is_hot
rcv_from_queue(1,10)=room_is_cool
VS_toss(3)=1
VS_toss(3)=3
VS_toss(3)=0
VS_toss(3)=2
(initial state)
send_to_queue(1,10,close_door)
send_to_queue(1,10,open_door)
VeriSoft Simulator
Dismiss
VeriSoft Simulator – Process 2
Step
Next
Continue
Print
Quit
VeriSoft Simulator – Process 1
Step
Next
Continue
Print
Quit
VeriSoft Simulator – State Space Filter
Text Regular Expression:
Match
Clear
Quit
AND
Labels
Processes
send_to_queue(1,10,room_is_hot)
send_to_queue(1,10,room_is_cool)
send_to_queue(1,10,open_door)
send_to_queue(1,10,close_door)
rcv_from_queue(1,10)=room_is_hot
rcv_from_queue(1,10)=room_is_coo
1
2
pts/2 /home/god/verisoft/examples/ac-controller

Move the pointer over a node to see which state it represents.

Assertion violation!

```
{
  ch
  me
  wh
```

```
      is_door_closed=1;
      if (is_room_hot)
        ac=1;
    };

    /* test */
    if (is_room_hot && is_door_closed)
      VS_assert(ac);

  };

}

void Environme
{
```

```
comeback $ verisoft main.c -simul error1.path
gcc -I/home/god/verisoft/bin /home/god/verisoft/bin/verisoft_simul_SunOS_5.5.1.o -
DVERIFY -g main.c
/home/god/verisoft/bin/simul.tcl error1.path
Loading sss.VS for state space view (please wait)...
Done.
```

Figure 7: Screenshot of VeriSoft in interactive simulation mode

releases of the call-processing software running on Lucent's CDMA base-stations, a multi-billion dollar product line. These projects range from so-called *"white-box" unit testing* of small critical applications described by at most a few thousands of lines of code, as in the case of [GHJ98], to *"black-box" system testing* of huge applications involving many concurrent components implemented by millions of lines of code, as in the case of [CGP02]. In each case, significant previously-unknown bugs were revealed through the use of VeriSoft.

The fact that VeriSoft detected errors that escaped traditional testing is not surprising, considering the following factors:

- Complex concurrent reactive systems are notorious for exhibiting a very large number of different behaviors.
- Traditional testing is of limited help since test coverage is bound to be only a minute fraction of all possible behaviors of the system.
- Systematic state-space exploration can expose previously unknown bugs by exercising the system under test in enormously more possible ways.

A key strength of the VeriSoft approach compared to other types of model checking is that it is extremely *versatile*: VeriSoft does not rely on any specific assumption about the code representing the behavior of processes, which can be written in any language and does not even need to be available.[3] Applying VeriSoft to a program written in a new language merely requires integrating VeriSoft into a test environment for that language, as described later in this section.

A second key strength of the VeriSoft approach is that it is *scalable*, in the following sense: the applicability of VeriSoft depends on the amount of nondeterminism in the system being analyzed, not on the size of the code describing the application itself. Nondeterminism comes from either concurrency or explicit nondeterminism introduced by the user using the special VS_toss operation. In the case of tens of concurrent processes, the amount of nondeterminism due to concurrency in the system is typically too large for VeriSoft to be very effective. An alternative is then to deliberately "hide" some system processes and communication objects from VeriSoft in order to reduce the amount of nondeterminism visible to the tool and hence reduce the size of the state space being explored. Since the user has the responsibility of declaring what operations are visible to VeriSoft anyway, omitting some processes or communication objects is easy. For instance, an entire communication switch can be be viewed as a huge black-box and multiple concurrent test-drivers controlled by VeriSoft can simulate various sequences of external events occurring at different interfaces of the switch (simulating other switches, user traffic and hardware failures, for example); even though VeriSoft does not control the nondeterminism (if any) inside the black-box itself with this approach, this can still be a very challenging test for the application. This type of approximation is necessary for analyzing applications of the size and complexity of the CDMA call-processing software [CGP02]. Obviously, if all the sources of nondeterminism are not under the control of VeriSoft, completeness of verification results and reproducibility of error traces cannot be guaranteed anymore.

[3] In this general discussion, the term VeriSoft is meant to denote both the current implementation of the tool and the VeriSoft approach in general, eventhough the former is obviously more limited than the latter.

We now discuss other practical considerations associated with this approach to software model checking.

Test automation. Using VeriSoft for testing the correctness of a software product requires *test automation*, i.e., the ability to run and evaluate tests automatically. As testers know, developing a testing infrastructure that provides test automation can be in itself a significant effort. When test automation is already available, starting to take advantage of VeriSoft to significantly increase test coverage is usually easy since it may just involve modifying existing test scripts into nondeterministic ones and/or running multiple test scripts in parallel under the control of the VeriSoft scheduler.

Integration into testing environment. VeriSoft needs to be integrated into the execution environment of the system under test so that it can control at run-time the execution of system processes. The primary task involved here is to declare which system calls of which processes are to be intercepted by VeriSoft and viewed as visible operations. Minimally, visible operations may simply include operations such as VS_toss and VS_assert for "black-box" testing of large applications. In the case of unit testing of applications containing only a handful of processes, system calls related to communication can also be declared as visible by mapping these to corresponding operations included in built-in VeriSoft libraries; for instance, sending a message (using whatever protocol is used by the application) can be mapped to a VeriSoft `send_to_queue` operation. Note that the actual system/protocol call need not be replaced by `send_to_queue`, it can just be annotated with the occurrence of that event. Mapping system calls related to communication to operations understood by VeriSoft can be tricky when complicated and unusual communication objects are used. Instrumenting the execution itself can be done by overriding system calls at compile/link time, or via a binary-code or OS-kernel instrumentation, or through the use of wrap-up functions intercepting events going in and out of the application being tested.

Test drivers. Like most model checkers, VeriSoft requires an executable representation of the environment of the system being analyzed in order to drive its executions. Thanks to the VS_toss operation supported by VeriSoft, nondeterministic programs can be used as environment models (test drivers). Nondeterminism makes it possible to write very compact and elegant programs for generating large numbers of sequences of input events (test scenarios). Since the size of the state space depends on the amount of nondeterminism in the system, VS_toss should be used with care.

Specifying properties. Although VeriSoft can simply be used to detect standard errors such as deadlocks and segmentation faults, it is preferable to specify application-specific properties by means of assertions in test drivers in order to check the functional correctness of the software application. Obviously, assertions previously inserted in the code itself by application developers can also be tested. Another possibility is to use tools (like Purify) that automatically insert assertions to check for standard programming errors such as memory leaks.

State explosion. The main practical limitation of VeriSoft, and of model checking in general, is the *state-explosion problem*: it is very easy for the user to define a state space that is too large to be explored exhaustively. State explosion can be controlled by limiting the amount of nondeterminism visible to VeriSoft, as discussed above. However, hiding nondeterminism due to concurrency inside the application being tested may result in errors being missed.

7 Comparison with Related Work

Essentially two approaches to software model checking have been proposed and are still actively being investigated. The first approach is the one presented in the previous sections and originally introduced in [God97]. The second approach consists of automatically extracting a model out of a software application by statically analyzing its code and abstracting away details, applying traditional model checking to analyze this abstract model, and then mapping abstract counter-examples (if any) back to the code. The investigation of this second approach can be traced back to early attempts to analyze concurrent programs written in concurrent programming languages such as Ada (e.g., [Tay83, LC91, MR93, Cor96]). Other relevant work includes static analyses geared towards analyzing communication patterns in concurrent programs (e.g., [Col95, Cri95, Ven97]). Recently, several efforts have started aiming at providing model-checking tools based on source-code abstraction for mainstream popular programming languages such as C and Java. For instance, Bandera [CDH$^+$00] can translate Java programs to the (finite-state) input languages of existing model checkers like SMV and SPIN, using user-guided abstraction, slicing and abstract interpretation techniques. SLAM [BR01] can translate sequential C programs to "boolean programs", which are essentially inter-procedural control-flow graphs extended with boolean variables, using an iterative automatic abstraction-refinement process based on the use of predicate abstraction and a specialized model-checking procedure. Feaver [HS99] can translate C programs into Promela, the input language of the SPIN model checker, using user-specified abstraction rules. Java PathFinder [VHBP00] can perform model checking of multi-threaded Java programs using a blend of static and dynamic program-analysis techniques. For the specific classes of concurrent programs that these tools can handle, the use of abstraction techniques can produce a "conservative" model of a program that preserves basic information about the execution and communication patterns taking place in the system executing the program. Analyzing such a model using standard model-checking techniques can then prove the absence of certain types of errors in the system, without ever executing the program itself. In contrast, VeriSoft is based on the *dynamic* observation of the "actual" processes of the concurrent system. This makes possible a much closer examination of the behaviors of the system, and the detection of subtle implementation errors that would typically be missed in practice using abstraction techniques, because either the abstraction does not preserve all the details relevant to that particular error, or because it would be hidden in a multitude of higher-level warnings. Moreover, VeriSoft does not rely on any specific assumption about the static structure of the programs used to represent the behavior of processes, which can actually be written in any language, or even be unavailable. In summary, the dynamic (VeriSoft) and static (abstraction-based) approaches to software model checking inherit the well-known advantages and limitations of, respectively, dynamic and static program analysis, and are therefore complementary.

The closest alternative to the type of software model checking developed in our work is perhaps specification-based testing frameworks for reactive programs (e.g., [YL91, DY94, Ric94, CRS96, JPP$^+$97]). Given a specification of the input/output behavior of the system being tested represented by a finite-state machine (or a product of finite-state machines [FJJV96]) expressed in some modeling language, these techniques and tools can automatically generate a set of test sequences that cover the specification according to various coverage criteria. In contrast, VeriSoft generates test scenarios *dynamically* at run-time: state-space exploration is performed while the system is executing, and the outcome of previous test sequences (i.e., paths in the state space) typically influences the generation of following test sequences (by the use of partial-order reduction methods). Moreover, using VeriSoft does not require a specification of the input/output behavior of the system under test written in some specific FSM modeling language; instead, the environment of an open system can be represented by one or several processes executing arbitrary code, and the joint behavior of all these processes is then checked for "global" properties when exploring the resulting state space, in the style of what is usually done with model checking.

Another related and complementary area of research concerns the design of debuggers for distributed and parallel programs (e.g., [CMN91]). These tools are used to monitor the execution of concurrent processes running in their actual environment. Work in this area discuss techniques for, among others, (1) instrumenting the execution of processes while minimizing the impact of the instrumentation on the timing (scheduling) between the different processes, (2) storing a minimum amount of information for faithfully replaying ("roll-back") very long scenarios leading to errors, and (3) obtaining a consistent representation of a state ("snapshot") of a distributed/concurrent system. These problems are avoided with our approach since (1) all the sources of nondeterminism are fully controlled by the VeriSoft scheduler, (2) the purpose of our approach is to make possible the systematic analysis of short executions of a concurrent system, rather than analyzing very long ones (e.g., containing millions of process transitions), and (3) our analysis is performed by examining only the global states of the concurrent system, which the scheduler can easily re-create. Note that VeriSoft does not generally preserve quantitative properties (related to timing, performance, etc.) of the whole concurrent system.

Other complementary work includes tools (like Purify) that automatically instrument code or executable files for monitoring program executions and detecting at run-time standard programming and memory-management errors such as array out-of-bounds and memory leaks. Also, several tools for monitoring at run-time the behavior of a reactive program and comparing this behavior against an application-specific high-level specification (typically a finite-state automaton or a temporal-logic formula) have recently been developed (e.g., [Dru00, HR01]). These tools can be used in conjunction with VeriSoft in order to increase the likelihood of errors being detected.

We conclude this section by mentioning other work that either makes use of VeriSoft or was directly inspired by it. [BG97] describes how to automatically synthesize a finite-state machine that simulates all the sequences of visible operations of a given process that were observed during a state-space exploration performed by VeriSoft. [CGJ98] discusses algorithms for automatically closing an open concurrent reactive C program with its most general environment, i.e., the environment that can provide any input at any time to the program; the result is a nondeterministic closed (i.e., self-executable) program that can exhibit all the possible behaviors of the original program and can be analyzed with VeriSoft. [God99] studies how to adapt model-checking symmetry reduction methods to make these compatible with a state-less search, i.e., without relying on explicit encodings of system states. [Sto00] discusses how to optimize a (state-less) search in the state space of a multi-threaded program whose threads use locks to protect access to shared data structures. Applications of VeriSoft to analyze programs specified using domain-specific primitives implemented as Java libraries are described in [GJJL00] and [GHJL00]. [GK02] discusses how to exploit heuristics using genetic algorithms to guide a search towards error states in very large state spaces; this framework has been implemented in conjunction with VeriSoft and evaluated with several examples of C programs. [BFG02] presents an overview of a tool for automatically discovering and systematically exploring Web-site execution paths; this tool includes VeriSoft as one of its components.

8 Conclusions

We have presented a new search technique for efficiently exploring the state space of a concurrent system composed of processes described by programs written in full-fledged programming languages such as C or C++. For finite acyclic state spaces, we showed that our algorithm can be used for detecting deadlocks and assertion violations without incurring the risk of any incompleteness in the verification results. In practice, our algorithm can be used for systematically and efficiently testing the correctness of any concurrent system, whether its state space is acyclic or not. This algorithm is built upon existing state-space pruning techniques known as partial-order methods [God96]. It extends the scope of verification by state-space exploration from modeling languages to programming languages.

This algorithm has been implemented in VeriSoft. Like a traditional model checker explores the state space of a system modeled as the product of concurrent finite-state components, VeriSoft systematically explores the "product" of concurrent (Unix-like) processes by using a run-time scheduler for driving the entire application through the states and transitions of its state space. It thus adapts model checking into a form of systematic testing that simulates the effect of model checking while being applicable to concurrent processes executing arbitrary code.

Two features distinguish VeriSoft from every other "model checker" (i.e., systematic state-space exploration tool) we are aware of: (1) VeriSoft is the first model checker that does not require the use of any specific modeling or programming language; (2) VeriSoft is the first model checker that does not compute and store representations of visited system states and performs a state-less search instead.

Since made publicly available in 1999, VeriSoft has been licensed to hundreds of users in industry and academia. Inside Lucent Technologies, it was applied successfully to analyze several software products in various business units and application domains (switch maintenance, call processing, network management, etc.). Because VeriSoft can automatically generate, execute and evaluate thousands of tests per minute, it can quickly reveal behaviors that are virtually impossible to detect using conventional testing techniques.

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VeriSoft can be downloaded from `http://www.bell-labs.com/projects/verisoft`.

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A Correctness Proofs

Theorem 1 *Consider a concurrent system as defined in Section 2, and let A_G denote its state space. Then, all the deadlocks that are reachable after the initialization of the system are global states, and are therefore in A_G. Moreover, if there exists a state reachable after the initialization of the system where an assertion is violated, then there exists a global state in A_G where the same assertion is violated.*

Proof:

By definition, a deadlock is a state where the execution of the next operation of every process in the system is blocking. Since we assumed that only executions of visible operations may be blocking, all deadlocks are global states.

Let s be a reachable state where an assertion a is violated. Let P_i be the process containing the assertion a. We know that the next operation to be executed by P_i in s is the assertion a, which is a visible operation. For every process P_j other than P_i, let o_j denote the next visible operation that process P_j will eventually execute from s. Consider the global state s' where, for all processes P_j, $j \neq i$, the next operation to be executed by P_j is the visible operation o_j, and the next operation of process P_i is the assertion a. Clearly, the global state s' is reachable from state s. Moreover, since only invisible operations may have been executed from s to s', assertion a is still violated in s'. (The execution of invisible operations in a process may not change the value of any variable or data structure local to another process.) Finally, since s' is reachable from s which is itself reachable after the initialization of the system, there exists a concurrent execution of the system that reaches the global state s' after the initialization of the system. Any sequence w of process transitions such that the sequence of visible operations in w can be observed during the concurrent execution leading to s' defines a path from s_0 to s' in the global state space A_G of the system. Therefore, s' is in A_G. ■

Let us now turn to the proof of Theorem 2. To establish this result, we use the notion of Mazurkiewicz's traces [Maz86]. *Traces* are defined as equivalence classes of sequences of transitions. Given a set $\mathcal{T}$ and a valid dependency relation $D \subseteq \mathcal{T} \times \mathcal{T}$ as defined in Definition 1, two sequences over $\mathcal{T}$ belong to the same trace with respect to D (are in the same equivalence class) if they can be obtained from each other by successively exchanging adjacent transitions which are independent according to D. For instance, if t_1 and t_2 are two transitions of $\mathcal{T}$ which are independent according to D, the sequences $t_1 t_2$ and $t_2 t_1$ belong to the same trace. A trace is represented by one of its elements enclosed within brackets and, when necessary, subscripted by the alphabet $\mathcal{T}$ and the dependency relation. Thus the trace containing both $t_1 t_2$ and $t_2 t_1$ could be represented by $[t_1 t_2]_{(\mathcal{T},D)}$. A trace corresponds to a partial ordering of symbol occurrences and contains all linearizations of this partial order. If two independent symbols occur next to each other in a sequence of a trace, the order of their occurrence is irrelevant since they occur concurrently in the partial order corresponding to that trace. Moreover, all sequences of transitions in a trace lead to the same state if executed from the same starting state. The latter property is formalized as follows.

Theorem 3.10 of [God96] *Let s be a state in A_G. If $s \overset{w_1}{\Rightarrow} s_1$ and $s \overset{w_2}{\Rightarrow} s_2$ in A_G, and if $[w_1] = [w_2]$,* *then $s_1 = s_2$.*

We will also make use of the two following lemmas from [God96]. These two lemmas state basic properties of persistent sets.

Lemma 4.2 of [God96] *Let s be a state in A_G, and let d be a deadlock reachable from s in A_G by a nonempty sequence w of transitions. For all $w_i \in [w]$, let t_i denote the first transition of w_i. Let Persistent_Set(s) be a nonempty persistent set in s. Then, at least one of the transitions t_i is in Persistent_Set(s).*

Lemma 6.8 of [God96] *Let s be a state in A_G, and let w be a nonempty sequence of transitions from s in A_G. For all $w_i \in [w]$ from s in A_G, let t_i denote the first transition of w_i. Let Persistent_Set(s) be a nonempty persistent set in s. If none of the t_i are in Persistent_Set(s), then all the transitions in Persistent_Set(s) are independent with all the transitions in w.*

To establish the correctness of Algorithm 2, we first prove the following lemma. Assume that all that concerns sleep sets in Algorithm 2 is not implemented (or equivalently that the sleep set associated with every global state reached during the search is empty). We now prove that, under this assumption, if there exists a sequence of transitions in A_G from s_0 to a deadlock or to a state s where an assertion a is violated, then Algorithm 2 without using sleep sets will eventually visit this deadlock or a state where the assertion a is violated, provided that A_G is finite and acyclic.

Lemma 1 *Consider a concurrent system as defined in Section 2, and let A_G denote its state space. Assume A_G is finite and acyclic. Let A_R be the state space explored by Algorithm 2 without using sleep sets. Let s be a state in A_R. Let d be a deadlock reachable from s in A_G by a sequence w of transitions. Then, d is also reachable from s in A_R. Moreover, if s' is a state where an assertion a is violated that is reachable from s in A_G by a sequence w' of transitions, then there exists a state (not necessarily s') reachable from s in A_R where the assertion a is violated.*

Proof:

The proof proceeds by induction on the length of w and w'. For $|w| = 0$ and $|w'| = 0$, the result is immediate. Now, assume the theorem holds for paths (sequences of transitions) of length $n \geq 0$ and let us prove that it holds for paths of length $n + 1$.

Assume a deadlock d can be reached from s by a path w of length $n+1$ in A_G. For all $w_i \in [w]$, let t_i denote the first transition of w_i. By Theorem 3.10 of [God96], we know that, for all $w_i \in [w]$, $s \overset{w_i}{\Rightarrow} d$. Let Persistent_Set($s$) be the nonempty persistent set that is selected in s by Algorithm 2, i.e., the set of transitions that are explored from s in A_R. By Lemma 4.2 of [God96], we know that at least one of the transitions t_i is in Persistent_Set(s). Since t_i is in Persistent_Set(s), it is explored from state s, and thus a state is reached in A_R from which a path of length n leads to the deadlock d. This together with the inductive hypothesis proves the lemma for the deadlock case.

We now consider the case of an assertion violation. Assume that a state s' where an assertion a is violated can be reached from s by a path w' of length $n + 1$ in A_G. Let Persistent_Set(s) be the nonempty persistent set that is selected in s by Algorithm 2, i.e., the set of transitions that are explored from s in A_R. For all $w'_i \in [w']$, let t'_i denote the first transition of w'_i. By Theorem 3.10 of [God96], we know that, for all $w'_i \in [w']$, $s \overset{w'_i}{\Rightarrow} s'$. If at least one of the transitions t'_i is in Persistent_Set(s), it is explored from state s, and thus a state is reached in A_R from which a path of length n leads to s'.

Otherwise, by applying Lemma 6.8 of [God96] to s and w', we know that all the transitions in Persistent_Set(s) are independent with all the transitions in w'. Consequently, for every state s_j reached after executing one transition in Persistent_Set(s) in A_R, the sequence of transition w' is still executable from s_j in A_G and leads to a state s'_j where the assertion a is violated (this follows from Definition 1). By applying the same reasoning to any state s_j and since all the executions of the system are finite (since its state space is finite and acyclic), one concludes that a transition t'_i is eventually executed from a successor state s_k of s such that all the transitions from s to s_k are independent with all the transitions in w'. After the execution of t'_i from s_k, a state s_l is reached in A_R from which a path of length n in A_G leads to a state where the assertion a is violated. This together with the inductive hypothesis proves the lemma for the case of an assertion violation. ■

From Lemma 1 it is then immediate to conclude that a state-less search using only persistent sets and started in the initial state of A_G will detect all the deadlocks and assertion violations in A_G. We now show that the use of sleep sets as described in Algorithm 2 preserves this result.

Theorem 2 *Consider a concurrent system as defined in Section 2, and let A_G denote its state space. Assume A_G is finite and acyclic. Then, all the deadlocks in A_G are visited by Algorithm 2. Moreover, if there exists a global state in A_G where an assertion is violated, then there exists a global state visited by Algorithm 2 where the same assertion is violated.*

Proof:

Consider a deadlock d or a state s' where an assertion is violated that is reachable from the initial global state s_0. Imagine that we fix the order in which transitions selected in a given state are explored and that we first run Algorithm 2 without sleep sets. Let A_R be the state space explored during this run. Assume that, for every state s in A_R, the transitions explored from s are sorted from left to right following the order in which they are explored: t_1 is to the left of t_2 if t_1 is explored before t_2. Then, we run Algorithm 2 with sleep sets while still exploring transitions in the same order. The important point is that the order used in both runs is the same, the exact order used is irrelevant. By Lemma 1, we know that, if d is a deadlock, d is visited by Algorithm 2 without sleep sets, while if an assertion a is violated in s', a state s'' where the same assertion is violated is visited by Algorithm 2 without sleep sets. We now prove that the leftmost path in A_R leading to d or to a state where the assertion a is violated is still explored in the second run when using Algorithm 2 with sleep sets.

Let $p = s_0 \overset{t_0}{\rightarrow} s_1 \overset{t_1}{\rightarrow} s_2 \ldots s_{n-1} \overset{t_{n-1}}{\rightarrow} s$ be this path. The only reason why it might not be fully explored (i.e., until s is reached) by the algorithm using sleep sets is that some transition t_i of p is not taken because it is in the sleep set associated with s_i. This means that t_i has been added to the sleep set associated with some previous state of the path p and then passed along p until s_i. Let us prove that this is impossible.

Assume that t_i is in the sleep set associated with state s_i, denoted $s_i.Sleep$, during the exploration of the path p. Hence, t_i has been added to the sleep set associated with some previous state s_j, $j < i$, of the path p and passed in the sleep set associated with the successor states of s_j along the path p until s_i. Formally, $t_i \notin s_j.Sleep$ when s_j is visited along this path and $t_i \in s_k.Sleep$ for all states s_k, $j < k \leq i$. This implies that t_i has been explored *before* t_j from s_j since a transition is introduced in the sleep set after it has been explored (line 14 of Algorithm 2). Moreover, all transitions that occur between t_j and t_i in p, i.e., all t_k such that $j \leq k < i$, are independent with respect to t_i. Indeed, if this was not the case, t_i would not be in $s_i.Sleep$ since transitions that are dependent with the transition taken are removed from the sleep set (line 11 of Algorithm 2).

Consequently, $t_i t_j \ldots t_{i-1}$ (the sequence $t_j \ldots t_{i-1} t_i$ where t_i has been moved to the first position) is in $[t_j \ldots t_{i-1} t_i]$. Thus, $t_i t_j \ldots t_{i-1}$ and $t_j \ldots t_{i-1} t_i$ are two interleavings of a single trace, and hence lead to the same state: $s_j \overset{t_i t_j \ldots t_{i-1}}{\Rightarrow} s_{i+1}$. Since there is a path $s_j \overset{t_i t_j \ldots t_{i-1}}{\Rightarrow} s_{i+1}$ from s_j, and since t_i is explored before t_j in s_j, the application of Lemma 1 to the state reached after the execution of t_i from s_j implies that the path p is not the leftmost path in A_R leading to d or to a state where the assertion a is violated. A contradiction. ■

Finally, it is worth noticing that all the above results also hold when a valid *conditional* dependency relation is used. Moreover, in that case, the above results hold without requiring the valid conditional dependency relation to be *weakly uniform* [God96].