MATISSE: Methodologies and Technologies for Industrial Strength Systems Engineering

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Practitioners Handbook

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1 Introduction

The core objective of the MATISSE project is the development of industrial-strength methodologies and associated technologies for the engineering of software-based critical systems. These methodologies and technologies will support industry in providing essential services that are highly dependable and therefore lead to increased public confidence and trust in these services.

Over the last 30 years, computer scientists have been developing and advocating the use of rigorous mathematically-based software engineering techniques, so-called *formal methods*, that support validation throughout the development life-cycle by providing rigorous specification and design notations as well as proof techniques, model-checking techniques and simulation techniques. Formal methods also allow the complexity of systems to be dealt with through abstraction and modularity. Despite this, the use of formal methods is not widespread in industry because of various managerial, sociological and technological barriers.

This document proposes work practices and guidelines to software practitioners to be used through the lifetime of a project. The guidelines presented in this handbook were obtained from experience gained in developing the project case studies. The development of the case studies allowed us to identify and understand the most effective way of introducing and using formal methods in a range of industrial environments and to identify, understand and overcome the barriers to their uptake and propose enhancements to the existing tools used in the project.

Within the MATISSE project we consider techniques suitable for overall system design rather than for isolated pieces of software. The overall trajectory of system design generally ranges from *user needs*, representing the first informal ideas about the system to be designed, to a *real system* implementation in a specific environment of use. The chapters of this handbook follow the trajectory of a typical development lifecycle, so each chapter corresponds to a phase of the lifecycle:

- modelling notations;
- constructing models;
- analysing models;
- refining models; and
- implementing models.

1.1 B Development Life Cycle

1.1.1 Overview

The aim of our approach is to build a B [Abrial96] application from an initial informal statement which is enriched, structured, and formalized progressively. This demands that, while preserving its own nature and goals, each phase of the construction process is carried out according to the needs and constraints of the following phase. Thus, the results of the each phase can be

effectively reused by the next one, which optimizes the process and assures a good traceability across the phases. The role and the principles of the phases of the process are the following.

1.1.2 Comparison with Standard Development Life Cycle

Standard development life cycle (Figure 1.1) and formal development life cycle (Figure 1.2) are quite similar in their form, except that the former contains less steps: most unit, integration and validation tests[1] are suppressed as they are replaced by proof, performed during all the development process. On the other hand, formal specification phase is usually than its classical counterpart, as specification is verified in depth by proof during its writing.

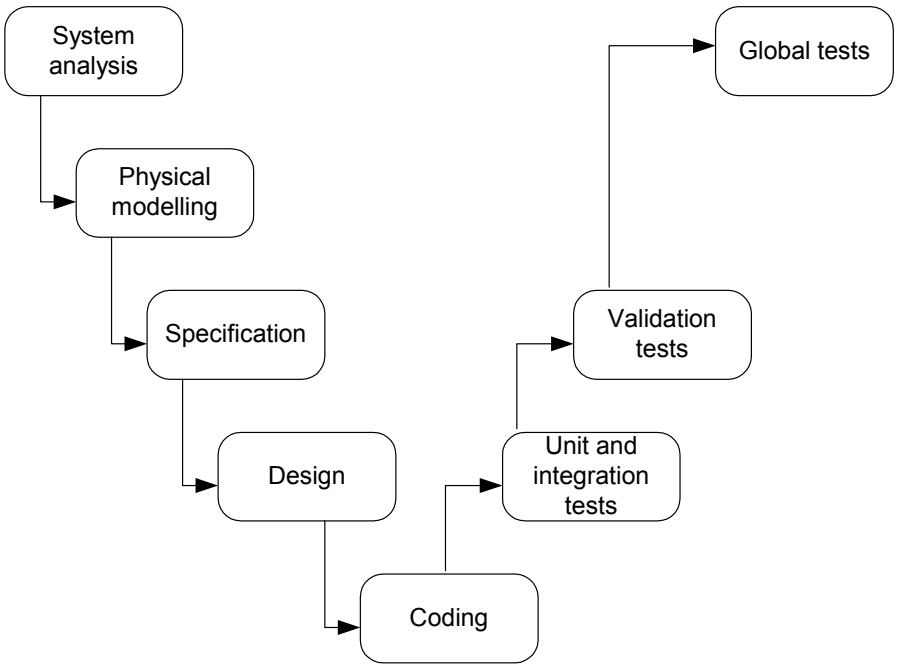


Figure 1.1: A classical V development life cycle.

The phases of the formal V lifecycle are:

- The requirements specification phase, that states the capabilities of the software to be developed and the operational and implementation constraints on it.

- The software specification phase, that defines abstractly which functions the software must perform in order to have the desired capabilities.

- The design phase, that defines the program implementing the software functions, and satisfying it's operational and implementation constraints. The design phase is divided in two sub-phases:

[1] Integration of third party libraries or non formally developed software should there be tested.

- The preliminary design phase, that defines the program components (sub-programs), their interfaces, and their organization.
- The detailed design phase, that completes the sub-programs definition, and produces the representation of the eventual executable program.
- The programming phase, that translates the program representation in the implementation programming language, and produces the actual executable program.
- The tests phase, that is divided in three sub-phases:
  - The unit tests phase, that checks each procedure of each sub-program.
  - The integration tests phase, that checks the cooperation of sub-programs.
  - The validation tests phase, that checks the adequacy of the program with respect to requirements.

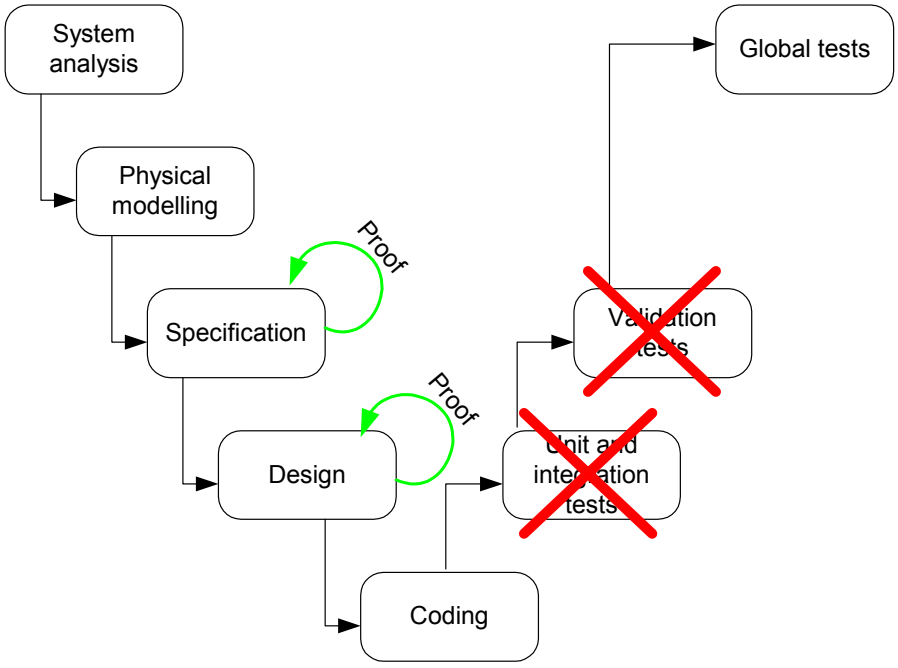


Figure 1.2: A typical B development life cycle.

1.1.3 Process Description

1.1.3.1 Requirements Specification

The requirements specification phase poses the problem to be solved by the software to be developed. Although informal, this first statement classifies and orders the requirements according to their nature and their safe level.

1.1.3.2 Software Specification

The software specification phase analyses the problem, and defines an abstract solution that meets requirements. The solution is defined in terms of functions and data, hierarchically organized from the most global and abstract description to the most detailed and concrete. The result of this phase is a semiformal model made of a functional model, a data model, and a collection of formal comments. The models describe the dependencies, causalities, and decompositions of functions and data. The formal comments are mathematical statements modelling functional requirements, safety requirements, and integrity properties of data of the model.

1.1.3.3 Preliminary Design

Preliminary Design transforms the semiformal model produced so far in a partial AMN model. This model introduces all the software components (abstract machines), defines their interfaces (constants, variables, and operations), and sets the implementation architecture (imports relations). The transformation principles are the following:

- Functions are transformed into AMN operations.

- Data is transformed into AMN variables or operation parameters.

- Levels of functional decomposition are transformed into layers of abstract machines: *i.e.*, the variables and the operations corresponding to data and functions of a decomposition level are grouped into one or several abstract machines. The criteria to group them are dictated by the constraints of AMN and the data model.

- Function decomposition is transformed into operation refinement: each operation associated with a decomposed function is implemented with the operations associated with the decomposition functions. In other words, the abstract machines associated with the non terminal levels of functional decomposition are implemented with the abstract machines associated with the next lower level of functional decomposition.

- Formal statements are transformed into invariants or operation definitions either in abstract machines or implementations. The result of preliminary design is then a layered architecture of abstract machines, where each layer represents a level of the functional decomposition, and is implemented, if it is not the bottom layer, with the abstract machines of the next lower layer. We shall see that in this kind of model, the safety and functional requirements are distributed at the different layers of abstract machines. Note that this model contains neither the refinements of abstract machines at intermediate layers, nor the refinements and the implementations of the abstract machines at the bottom - given implementations must be added to the partial AMN model.

1.1.3.4 Detailed Design

Detailed design finalizes the partial AMN model produced at the previous phase. It creates the refinements of the abstract machines of the intermediate layers and the refinements and the implementations of the abstract machines of the bottom layer which are developed in B. Therefore, the result of the detailed design is the complete AMN model of the software.

Some abstract machines are neither refined, nor implemented. They are supposed to formalize the interface of a component that is not formally developed, either because it is an already developed component or library, or because it is an interface with the underlying operating system, or because it is a low level component whose formal development would be too expensive compared with the expected benefits.

1.1.3.5 Programming

In the present approach the executable code of the software is obtained by translating, either automatically or manually, the implementations of the AMN model and, if necessary, by developing conventionally the code of the abstract machines that have not been formally refined and implemented.

1.1.3.6 Tests

Tests phase is divided into unit tests, integration tests, and validation tests. Unitary tests concern exclusively operations of abstract machines not formally refined and implemented and operations that could not be proved completely.

Integration tests concern the integration of formally and not formally developed modules. Functional tests are carried out as in conventional developments. Unitary, integration, and non intensive validation tests are performed by the development team. Then, the software is delivered to a validation team, independent from the development team, which performs intensive validation tests on the target computer.

1.1.3.7 Proof Validation and Program Proof

Proof validation is carried out at preliminary and detailed design phases by the validation team. It ensures that the specific proof rules added to discharge proof obligations generated by AMN modules are correct mathematical statements.

1.2 Case Studies

The driver for the research and development in MATISSE are three major industrial case studies representing a spectrum of essential services for the information society. The industrial case studies are:

- A railway signalling and control system;
- A diagnostic system for healthcare clinicians and researchers;
- An embedded verifier for a multi-applications smartcard system.

These three case studies are used to describe the guidelines and work practices through this handbook.

1.2.1 Transportation

1.2.1.1 Formal Methods in the French Railway Industry

The use of formal methods in the French railway industry was introduced by the SACEM system, an automatic train control system for RATP, the transit authority in Paris, in the beginning of 1980. The consortium of manufacturers in charge of this project decided to develop non-diverse software using a new technology to secure it, the Vital Coded Processor (VCP) technique, instead of diversified software using the concept of redundancy. The VCP technique consists in protecting each software information by a redundant code, checked at run-time. Because the software was not duplicated, a *zero default* design was required. Moreover, since validation by testing was considered insufficient by RATP, they asked manufacturers to use a new approach based on formal methods. The process was as follows:

1. The functional software requirements were re-written in a formal language.

2. The software source code, written in the language Modula 2, was completed with *pre-assertions* and *post-assertions*, and checked with a partially automated program proof activity.

3. A check between formal specification and code assertions was manually performed.

Despite the heaviness of this process and its weak automation, the confidence achieved by this first application had convinced RATP of the advantages of this approach. For the subsequent tender, the Meteor line, they required the use of formal method for safety critical software development. MATRA Transport International was chosen in 1992 to develop the automatic train operation system for the Meteor line, using the B method for the safety critical software.

1.2.1.2 Transportation Case Study

The railway application case for MATISSE concerns a generic product originally designed for the protection of trains. It is an Automatic Train Protection (ATP) transmission-based system using a movement authority principle.

The ATP gives the movement authority to trains to enter and travel through a specific section of track in a given direction. The movement authority used by the limitation of movement is located in fixed places. The ATP supervises and enforces the authority for movement to maintain a safe train separation and provides protection based on interlocking information to the train movement.

In addition to ATP, the Automatic Train Operation function (ATO) starts and stops the train, controls the train doors, and regulates the speed of the train as it travels over the railway network in accordance with the Automatic Train Supervision system inputs. The ATO maintains the train speed below the speed limits imposed by the ATP.

The intelligent part of ATP and the safety decision-making process is placed on-board the trains. As messages on a section are broadcast to all trains located in an area, the on-board equipment calculates the position of the train on the track and filters out the relevant data. The on-board equipment, programmed with the parameters of the train, checks the train movement against the safety constraints and, in automatic mode, achieves the ATO functions.

Description of the Automatic Train Control Functions

The main functions provided by the SACEM[2] Automatic Train Control (ATC) are:

- Checking train movement and passenger safety (ATP) consisting of:
    - Compliance with protection points (this ensures spacing, routing, end of track and floodgate protection),
    - Compliance with speed limits,
    - Rollback monitoring,
    - Run authorization by zone for passenger safety (Platform Emergency Stop and Platform Screen Doors status supervision), staff protection (KeySwitch) and train movement protection (Flank and Fouling protection).
    - Movement or speed conflicting with the above results in irreversible emergency braking until the train is completely stopped.
    - Authorisation to open the train and platform screen doors at the platform.
- Train operation:
    - In Manual mode:
        - Speed instructions delivered to the Train Operator complying with protection points, speed restrictions, and other conditions.
        - Control of the opening of Platform Screen Doors.
    - In Automatic mode (Automatic Train Operation):
        - Optimum train movement between stations, complying with protection points and arrival/departure timetables.
        - Stopping of trains in stations (the actual departure is controlled by the Train Operator).
        - Stopping of trains in change of end zones (according to the Automatic Train Supervision orders).
        - Control of the opening of Platform Screen Doors.

[2] The SACEM system is a automatic train control developed by Siemens Transportation Systems in the beginning of 1980.

- Control of the opening of Train Doors (only in Driverless Turnaround mode).
- Train start in automatic mode between stations (with no action by the Train Operator).

- Interface with Operations and Maintenance personnel:
  - SACEM ATC delivers reports used by Operation or Maintenance personnel.

1.2.2 Healthcare

PerkinElmer Life Sciences (later in the document referred to as Wallac) supplies complete analytical systems. The systems are used to provide researchers and clinicians with reliable determinations of substances found in blood or other biological sample materials that are difficult to measure. Typically, the substance to be measured is tagged with a radionuclide or a fluorescent or luminescent label. The amount of the substance is then determined from the *signal*, the amount of radiation or light given off by the label. The manufactured analytical systems include reagents, sample handling and measuring instrumentation, as well as computer software.

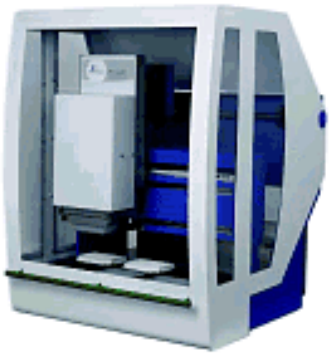


Figure 1.3: The Fillwell microplate liquid handling workstation.

The team of the Computer Science Department at Aabo Akademi University (later in the document referred to as Aabo Akademi) conducted a parallel development of a new product of Wallac's – Fillwell™, a workstation preparing samples [PerkinElmer01]. The workstation is shown in Figure 1.3. The system belongs to the class of products for drug discovery and bioresearch. The Fillwell microplate liquid handling workstation offers significantly advanced features in the line of the sample preparation systems. The Fillwell base unit consists of a dispense head dispensing liquid into microplates on a processing table. A gantry moves the dispense head with high precision and speed from one plate to another.

The Fillwell workstation is the first liquid handling system specifically designed for high density microplates. The system is modular and can therefore be customised into a variety of

configurations. The dispense head can have up to 384 tips attached. With the tips the head can perform automated pipetting into plates with 96, 384 or 1536 wells. The head provides a precise dispensing with volumes from 0.5 to 300 μl. The processing table contains up to 6 plate positions. An extension with three plate positions is easily removable. In Figure 1.3 this processing table extension has been removed. In order for the dispensing head to be able to reach all the positions on the processing table it is mounted on a gantry that can move in XYZ-directions. The precision of the gantry is very high with an accuracy of 100 μm. The system can function as a standalone workstation or be integrated into a robot.

The main application of the Fillwell workstation is drug discovery. Within this application area the system can be used for microplate replication, for dilution, transfer and addition of the liquid in the plates, for reformatting of plates with different densities (number of wells), as well as for rapid plate filling to homogenous and cell based systems. The FillWell system is a safety-critical system. Safety is the property of a system, which ensures that it will not harm humans, environment or equipment. The direct harm to the humans using the drug discovery systems is quite moderate according to the classification for normal safety-critical systems. However, the indirect harm caused by the results of incorrectly performed experiments might be catastrophic. Furthermore, the system can be used to handle extremely expensive substances (valued up to a billion EURO per kilogram) as well as serve as a part of an expensive production chain. Hence, failures of such a system might result in significant economical losses and the system can also be considered as money critical.

1.2.2.1 The Setting of the Case Study

This case study was organised using an industry-as-laboratory approach. This means that the researchers in academia provided the formal methods and B expertise while the experts at the R&D department of the industrial partner brought the domain knowledge to the team.

In order to come up with a good methodology for the development we started with a smaller part of the case study concentrating on modelling the dispense head and its movement up and down as well as its dispensing of liquid. However, the branch of development of medical and pharmaceutical equipment is highly competitive. Because of that, the requirement to preserve confidentiality of the equipment under construction was very strict. As the product was confidential for a long period of the project, we developed a mirror case study, called the Robot analyser, in order to have a public case study as an example of the methodology. We have chosen a mirror case study that models very well the significant and critical aspects of the Fillwell case study concerning the precision of movement and liquid handling. Both the case studies belongs to the same class of systems. The development of the Fillwell case study will be documented elsewhere [Bostroem03].

In the Fillwell case study we have a workstation dispensing to and aspirating from plates. Our mirror case study is a robot analyser that analyses samples on a plate that is placed on a movable operating table. The movement of the operating table corresponds to the movement of the dispense head, while the analysing corresponds to the dispensing. A first extension of the systems would be to take into consideration the XY-movement of the dispense head and, hence, the corresponding rotation of the robot to be able to work with more than one plate.

1.2.2.2 The Requirements of the Case Study

The Robot analyser consists of two interacting parts, a Robot and an Analyser. We concentrate here on the Analyser. The task of the Robot is to place a plate to be analysed into the Analyser, as well as to remove an analysed plate from it.

The Analyser has an operating table – a horizontal surface on which a plate is placed by the Robot. The Analyser can move its operating table to pick up a plate from the Robot, move the table to its analysing component, analyse the plate, as well as move the table to a position from which the Robot can remove the plate. During the analysis it is checked that there is enough liquid on the plate before the analysis and that the right amount of liquid was used during the analysis. The quality of the analysis should be kept at a constantly high level.

The Robot has two arms, one for placing a plate in the Analyser and the other for removing this plate. These arms are at different heights. Therefore, the table has to be at different positions for placing and removing plates:

- Upper end position: A plate is analysed by the analysing component.
- Middle position: Arm 1 of the Robot places a plate into the Analyser.
- Lower end position: Arm 2 of the Robot removes a plate from the Analyser.

Initially the Analyser is at its middle position and it is empty.

Due to the fact that the operating table and the Robot arms may collide we define *safety requirements* that the Robot analyser has to fulfil. These safety requirements are expressed in English. Later during the development they will be translated to predicates in B:

- When the Robot is ready to load the Analyser, arm 1 may extend only if the table is in its middle position.
- When the Robot is ready to unload the Analyser, arm 2 may extend only if the operating table is in its lowest position.
- The table may move only when the respective Robot arm has retracted.
- The table must not move beyond its upper and lower end position.

During a *typical course of events* the following sequence of actions happens with a plate. Here, we have not considered the possible failures.

1) The Analyser moves the operating table to the middle position (*move(zmid)*).
2) The Analyser waits to receive a plate from the first Robot arm after the Robot has picked up a plate and rotated to the Analyser (*receive*).
3) The Analyser moves the operating table to the upper end position (*move(zmax)*).

4) The Analyser analyses the plate (*analyse*).

5) The Analyser moves the operating table to the lower end position (*move(zmin)*).

6) The Analyser waits to deliver an analysed plate to the second Robot arm and the Robot rotates to pass the plate on to another device (*deliver*).

We can note that other courses of events are possible depending on the order in which the commands *move, receive, analyse* and *deliver* are given. In order to prohibit the Analyser from analysing a plate and moving an infinite number of times, only a certain number of moves are allowed to be performed in a row before some other action needs to be taken by the Analyser. Similarly, only a certain number of analyses on one plate are allowed before the plate has to be delivered.

1.2.3 Smart Card

The smart card case study was developed by Gemplus using the B method and the Atelier-B toolkit. This case study investigates how B can be used efficiently in the specific domain of the smart card. Gemplus has concentrated its effort on applying B at the software level, with the aim of producing more reliable and certifiable products.

Security in Java

Security is a critical aspect of downloadable code that originates from a remote source. It is not possible to know *a priori* the behaviour of the program. By executing the mobile code, we grant it the right to perform operations on our machine and we provide it access to our local resources: files, network access, local information and devices. Each piece of mobile code that is allowed to execute must be tightly controlled in order to ensure that it will not destroy our data, divulge confidential information, waste our resources, or use our machine as an intermediary to launch an attack on someone else. For the smart card it is important that an applet cannot have access to the data of other applets except by using the sharing mechanism, or accessing the code of the operating system.

The security in Java is partially based on language features like strong typing. By disallowing programmers to manipulate or forge pointers, Java controls the access a program can perform. But once the program is compiled in byte code, those properties can be violated by manipulating the byte code itself. The verifier is a key component of the Java security architecture. It examines incoming code in order to ensure that it is valid. It checks that the code respects the syntax of the byte code language and that it respects the language rules. Other components are responsible for protecting system resources from abuse but they depend on the verifier as they rely on language features such as access restrictions (`private`, `protected`, `final`, etc).

As opposed to ordinary native machine languages, the Java byte code language has been defined so that Java byte code programs can be statically verified and validated. The verifier is able to examine a compiled class and decide whether or not:

- it is syntactically correct and well-formed,
- it respects the language rules, and

- references from the current class to other classes are consistent.

The verifier performs this validation statically, and thus does not hinder the performance of the execution. Equivalent validation cannot be performed on programs compiled into native machine code, nor can it be performed on typical assembler source. It cannot even be performed on source code for many higher-level languages such as C or C++. Java byte code is probably the most feature-full language for which such verification is possible.

1.2.3.1 The Verifier

The verifier examines loaded classes to ensure that they are syntactically correct and well-formed, that they are consistent with other classes and that the byte code they contain respects the language rules. The verifier's purpose is to prevent the Java Virtual Machine (JVM) process from being subverted or corrupted by inconsistent or malformed CAP files[3], and to enforce the language's security mechanisms. The verifier checks the integrity of the constant pool and the syntax of all method and field declarations. It checks the consistency of the current class with its super classes (or super interfaces) and its implemented interfaces. For instance, it is checked that:

- references to the constant pool entries target entries of appropriate type for the referring entity,
- modifier flags (also called access flags) are consistent for the class and all its methods and fields,
- final classes are not subclasses and final methods are not overridden,
- classes really implement the interfaces they claim to implement,
- non-abstract classes do not have abstract methods.

The verifier also inspects the byte code for each method, checking that all opcodes and instruction operands are valid. For instance, it is checked that:

- branches target valid instructions,
- arrays of more than one dimension are not created.

The verifier keeps track of the types of all values on the stack and in the registers. It ensures that all instructions take appropriately typed arguments from the stack and registers. The property known as type-safety implies that:

- values of primitive types are never taken to be references, nor the reverse,

[3] The CAP file is produced by the converter and is the standard file format for the Java Card platform. This file contains the executable binary representation of the classes of a Java package.

- instances of a given class are used only where that class or a super class of that class is expected,
- an instance of a class can be used where a given interface is expected only if the class does implement the interface,
- objects are not used until they are initialised and that they are properly initialised.

Finally, whenever a reference to another class, method or field is about to be used by a running method, it is checked that the target exists, that it is accessible to the referring class and that it is used properly.

1.2.3.2 Java Card Verifier

Java and Java Card restriction

In the Java architecture, the class file verifier is used prior to execution. The code is loaded by the Java Virtual Machine, and then checked by the verifier before its execution by the interpreter. The Java verifier performs static checks and ensures some properties:

- there are no violations of memory management and no stack underflows or overflows,
- access restrictions are enforced,
- methods are called with appropriate arguments of the appropriate type,
- fields are modified with values of the appropriate type,
- objects are accessed with the appropriate type,
- no pointers are forged,
- no illegal data conversions are preformed, and
- binary compatibility rules are enforced.

However, the Java Card language differs from the Java language. First the execution architecture is different: the classes of an applet correspond to one or several packages. A converter is used in order to convert the classes of a package into a CAP file. In addition to the CAP file creation, an export file representing the public APIs of the package being converted is generated. Therefore, if one wants to verify a CAP file of a particular package, the CAP file, export file of the package and export files of the imported packages are needed as depicted in Figure 1.4. During verification, the verifier must ensure that the CAP file is internally consistent, consistent with the export files it imports and consistent with the export file that represents the API.

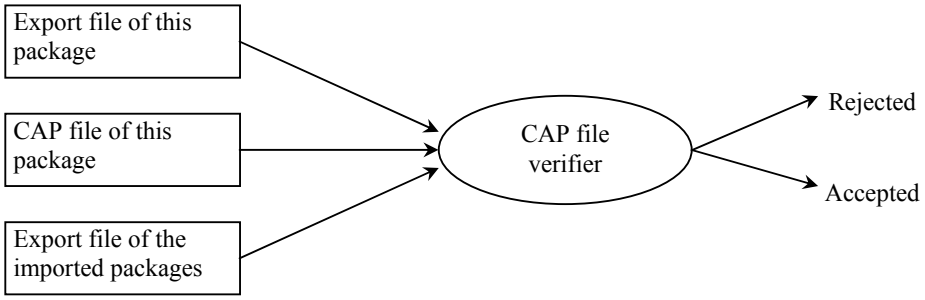


Figure 1.4: The CAP file verification.

Furthermore, the Java Card language is a subset of the Java language. In particular, several features of Java are not present: dynamic class loading, security manager, garbage collection and finalization, threads, cloning and access control in Java packages are not supported. The basic types `char`, `double`, `float` and `long` are not supported . Arrays of more than one dimension are also not supported by Java Card. These restrictions are imposed to check that:

- only supported data types are used and data of type `int` are used only if the Java Card interpreter supports them;
- no unsupported Java language features are used;
- usage of certain Java operations are within limited ranges; and
- no potential overflow or underflow can occur that might cause arithmetic results to be computed differently than they would be on the Java platform.

Most of these restrictions are checked by the converter and do not need to be explicitly performed by the verifier. In fact, if any data is of type `float`, the verifier would reject the program not because the type `float` is not supported but because the type is unknown. So the verification appears to be implicit. However, during verification, the CAP file verifier must ensure that a CAP file has the correct format. It also ensures that the byte codes within the CAP file fit to the set of structural constraints. Some of these constraints are the same as for Java and some others are particular to the CAP file structure such as the following:

- The package and each applet defined in the package must have a valid AID (Applet IDentifier) that is between 5 and 16 bytes long. The package AID and the applet AIDs must share the same RID (Resource IDentifier) number.
- An applet must define an `install` method with the correct signature so that instances of the applet can be appropriately created on the card.
- The order of class and interface definitions in a CAP file must follow the rules that interfaces appear before classes, and superclasses appear ahead of subclasses.
- The `int` flag is set if the `int` type is used in the CAP file.

2 Modelling Notations

In the MATISSE project, B is the core method; most of the other notations used either evolved from B or integrate B with an existing formal or informal notation. The exception is the CSP notation that will be used to complement B. This section briefly describes the notations that will be used through the handbook.

2.1 B Method

B is a formal specification method that allows, the properties required in a schedule of conditions to be rigorously expressed.

We can **prove in an automated way** that these properties are unambiguous, coherent and non-contradictory. Then, we **guarantee** by **mathematical proof** that these properties are satisfied in every stage of design.

This method and the associated proof that allow:

- to obtain technical specifications and system schedules of conditions that are **clear**, **structured**, **coherent** and **without ambiguity**,
- to develop software guaranteed by contract to be without defect.

The B method has been applied in large industrial projects? in fields such as real time, process controls, communications protocols, cryptographic protocols, embedded systems.

2.1.1 Synopsis

Created by Jean-Raymond Abrial, from the research of E.W. Dijkstra and C.A.R Hoare, the B method is described in [Abrial96]. It is based on the mathematical concept of set theory.

Generally, the initial expression of requirements is captured using natural language, or by combination of methods such as SADT or SA-RT or descriptions: such as graphs, automata, logic tables or Petri networks.

A B development begins with the construction of a model from the requirements, that describe

- the main state variables of the system,
- the properties (or invariant) which these variables will have to meet constantly, and
- the transformations of these variables by services (or operations).

The B model obtained constitutes a specification of what the system will have to realise (the *what*). The B model is then refined, *i.e.*, specialised, until a complete implementation of the software system is obtained (the *how*). Several refinements can fulfil a specification, the choice of solutions depending on various criteria such as the simplicity of proofs, and the architecture of the system. The refinement can also be used as a specification technique. In this case, the refinement enables progressive inclusion, of the problem's details in the formal development. The formal specification is then realised gradually rather than directly.

Using B in the development of a system is therefore about:

- Clearing all ambiguity right from the initial interpretation of the requirement,
- Constructing a specification that is both coherent and conforms to the requirement (the model),

- Elaborating the software system that realises the specification, in successive stages.

The coherence of the model and then the conformity of the final program to this model are guaranteed by mathematical proofs. The demonstration of these proofs can be carried out by using automatic proof tools, such as those provided by Atelier-B [AtelierB98].

2.1.1.1 Abstract Machines

The abstract machine is the basic mechanism of the B method. It is a concept very close to well known programming notions, such as modules, classes or abstract data types.

A machine contains variables and operations. It encapsulates variables. The operations enable access to the variables and their manipulation.

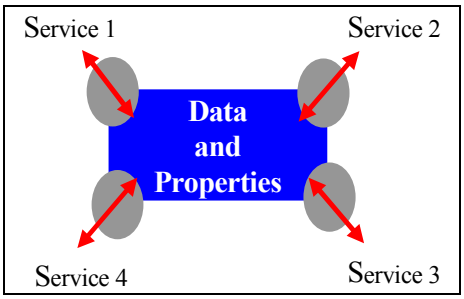


Figure 2.1: Abstract machine.

The elements of an abstract machine are described through expressions using mathematical concepts such as sets, relations, functions or sequences. The static laws, which the variables must obey, are defined by predicates and constitute the invariant of the abstract machine.

The specification of the behaviour of the operations uses a pseudo-code made of generalised substitutions. The substitutions are defined formally in the way they transform predicates. A given operation of an abstract machine can contain a pre-condition: it is a predicate expressing the conditions necessary to invoke the operation. An operation also contains an action: it is a substitution describing how the variables of a machine are manipulated. As an abstract machine is defined in the specification phase, some concrete substitutions such as sequencing or loops are forbidden. It is supposed to describe *what* the operation must do, not *how*. The parallel substitution, does not prescribe the application order of the individual substitutions. The nondeterminism present in the substitutions leave open choices for further development.

To ensure that an operation call preserves the static properties of the abstract machine (the invariant), a *Proof Obligation* has to be shown to hold. This Proof Obligation is automatically constructed from the formal definition of the substitutions in the abstract machine.

2.1.1.2 Refinements and Implementations

The mechanism of refinement consists of successively reformulating the variables and operations of the abstract machine, so that a computer program finally results. The intermediary stages of reformulation are called refinements, and the final level of refinement, the implementation.

The refinements must redefine the operations of the abstract machine. Each B component (abstract machine, refinement or implementation) is defined using the B Language.

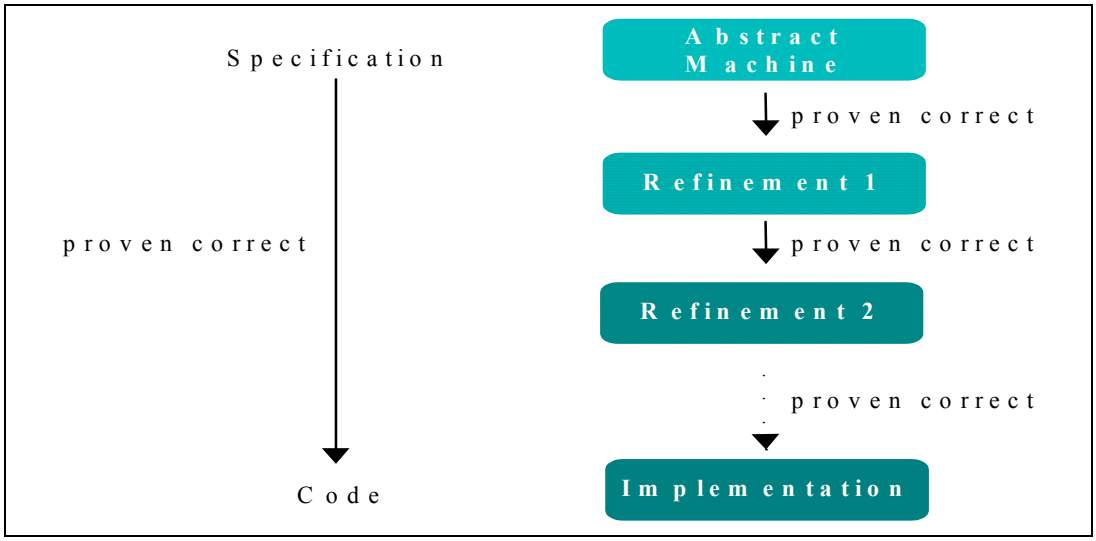


Figure 2.2: Refinement mechanism.

During each refinement, the behaviour of an operation has to be proven to be compatible with the operation it refines. This guarantees that the code of an implementation will conform to its specification in the abstract machine.

The production of a refinement:

- progressively transforms the abstract data types (sets, relations, functions, sequences) into concrete data types (scalar variables, arrays);
- gradually clears the level of non determinism of the substitutions; and
- replaces abstract substitutions (parallelism, choice) with concrete substitutions (sequencing, conditionals, loops).

2.1.1.3 Project

A B project enables the development of a piece of software realising a particular system. To realise a complete project, we use the mechanisms of composition and decomposition of abstract machines. As soon as the level of complexity of the refinement of an abstract machine is too high, we decompose it into several, more simple parts. The implementation (the final level of refinement of an abstract machine) can then be implemented on the specifications of one or several abstracts machines, which are themselves refinable. This is achieved with call to operations of the machines involved. The user of an abstract machine is therefore always an implementation. In this way, a project is constructed gradually, according to an architecture made from layers of abstraction.

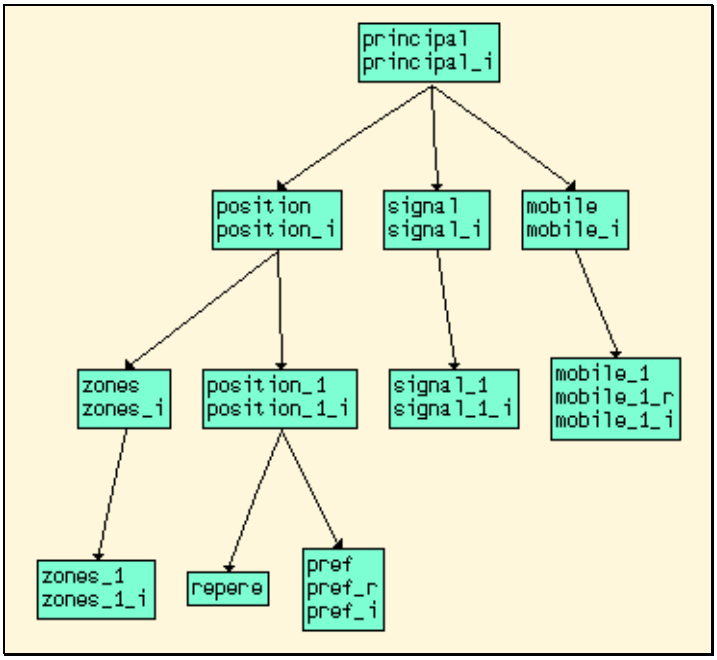


Figure 2.3: A B project.

Abstract machines of a lower level in the layer architecture of a project can pre-exist the rest of the development. Indeed, there is a library of abstract machines available. These machines encapsulate the more classic data structures, and are guaranteed, checked and proved.

2.1.2 The B Language

The B language consists of the following elements:

- Predicates
- Expressions
- Substitutions
- Components (abstract machines, refinements, implementations)

2.1.2.1 Predicates

The predicates used in the B language are a subset of the logic of first order predicates. They are formulae allowing the expression of properties relating to data.

Predicates are used directly in the B language to express properties relating to all the data of a component. For example, they express the invariant properties of variables and the pre-conditions under which an operation can be called (these conditions relate amongst others to input parameters of the operation).

They are used to express the *Proof Obligations*. The goal of a Proof Obligation is a predicate, which can either be proved or refuted. Hypotheses under which a proof is made are also predicates.

The predicates of the B language are:

- simple propositions (conjunctions, negation, disjunction, implication, equivalence);
- quantified predicates (universal, existential); and
- comparison predicates between expressions.

2.1.2.2 Expressions

An expression is a formula describing data. Every datum has a type and a value. The categories of expression are:

- basic expressions;
- boolean expressions;
- arithmetic expressions;
- maplet expressions;
- set expressions (empty set, set of integers, boolean,...);
- set constructions (set of sub-sets, union, intersection,...);
- relation expressions (identity, inverse, projection, composition, iteration, domain,...);
- function expressions (injections, surjections, bijections,...); and
- function constructions (constant functions, lambda expressions,...).

2.1.2.3 Substitutions

Generalised substitutions are mathematical notations that describe the transformation of predicates. They are used to describe the dynamic aspect of B components: their operations.

They describe the behaviour of operations at the abstract machine level as well as the refinement and implementation levels. The specification substitutions can be non-deterministic and non-executable, whereas the implementation substitutions correspond to instructions of a classic computing language.

They also enable the Proof of Obligations to be built automatically from the B components. For example, the Proof Obligation corresponding to the preservation of the invariant during the call of an operation is constructed by taking the invariant as hypothesis and the proof of substitution of the operation applied to the invariant as the goal.

The substitutions of the B language are:

- substitution becomes equal, substitution becomes such as, substitution becomes element of:

```
x := E;
```

- pre-condition substitution, to express pre-conditions of operation calls:

```
PRE G THEN S END;
```

- bounded choice substitution, SELECT substitution:

```
SELECT G THEN S END;
```

- substitutions ANY and LET which introduce data verifying certain properties:

```
ANY v WHERE P THEN S END,
LET v BE P IN S END;
```

- VAR substitution which introduces local variables:

```
VAR v IN S END;
```

- conditional IF and CASE substitutions:

```
IF P THEN S ELSE Q END,
CASE E OF
       EITHER l THEN S
       OR m THEN T
       …
       OR n THEN U END
END
```

- simultaneous substitution:

```
x, …, y := E, …, F,
x := E || … || y := F;
```

- loop substitution; and

2.1.2.4 Components

A B component can be an abstract machine, a refinement or an implementation. Components possess clauses incorporating the static and dynamic description of behaviour.

The main clauses are given in Table 1:

MACHINE	declaration of the abstract machine's name and the list of eventual parameters of the machine
REFINEMENT	declaration of the name of a refinement
IMPLEMENTATION	declaration of the name of an implementation
REFINES	declaration of the refined component
IMPORTS	when an implementation imports an abstract machine, it can use freely the latter's operations (but not its data)
SEES	when a component sees an abstract machine, it refers to this machine and can consult its data and use the operations which do not modify these data
INCLUDES	when an abstract machine includes another abstract machine it integrates the data of the included machine
DEFINITIONS	declaration of textual definitions which will be expanded in the component before any further analysis
CONSTRAINTS	declaration of the properties of the abstract machine's parameters
SETS	declaration of deferred and enumerated sets
CONCRETE_CONSTANTS	declaration of constants, concrete and implementable, which will be kept during successive refinements
ABSTRACT_CONSTANTS	declaration of abstract constants, which are non-implementable

	and must therefore be refined
PROPERTIES	declaration of the constants' properties
CONCRETE_VARIABLES	declaration of concrete variables, which are implementable and will be kept during the successive refinements
ABSTRACT_VARIABLES	declaration of the abstract variables, which are non-implementable and must therefore be refined
INVARIANT	declaration of invariant properties of the variables
INITIALISATION	initialisation of variables
OPERATIONS	declaration of the operations in the form of a header and body

Table 1: Clauses of a B component.

2.1.3 Examples

2.1.3.1 Modelling the Railway System

The railway example presented here [Butler02a] is a simplified version of the case study presented in Section 1.2.1. Figure 2.1 describes an abstraction of the railway states in B notation. `SECTION` and `TRAIN` are types used to identify track sections and trains respectively (1). A railway switch has two incoming and one outgoing sections (convergent) or vice-versa (divergent). Railway switches are regarded as special cases of track sections. The constant `switch` (2) represents the subset of sections that are switches, while the constant `net` is a graph representing the static connectivity between sections (*i.e.*, the connectivity regardless of the switch positions). In the *properties* clause, `switch` is declared to be a subset of `SECTION` (3), while `net` is declared to be a relation between `SECTIONS` (4). Here, we have omitted other properties of `net` such as the fact that a section cannot be connected to itself.

```
MACHINE  Railway
SETS  SECTION, TRAIN                                                    (1)
CONSTANTS  switch, net                                                  (2)
PROPERTIES
   switch ⊆ SECTION  ∧                                                  (3)
   net ∈ SECTION ↔ SECTION                                              (4)
VARIABLES  next, occp, front, back, braking, checked                    (5)
INVARIANT
   next ⊆ net ∧                                                         (6)
   next ∈ SECTION ⤖ SECTION ∧                                           (7)
   occp ∈ TRAIN ↔ SECTION ∧                                             (8)
   front ∈ TRAIN → SECTION ∧ front ⊆ occp ∧                             (9)
   back ∈ TRAIN → SECTION ∧ back ⊆ occp ∧                               (10)
   braking ∈ TRAIN → Bool ∧                                            (11)
   checked ∈ TRAIN → Bool ∧                                            (12)
   ∀ t1, t2 . (t1 ∈ TRAIN ∧ t2 ∈ TRAIN ∧
         front(t1) = back(t2) ⇒ braking(t1) = TRUE                      (13)
```

Figure 2.4 Variables and invariant of abstract railway.

The variables of the machine (5) are typed in the *invariant* clause. Variable `next` represents the dynamic connectivity between sections based on the positions of switches. Note that we have abstracted away from the actual positions of switches and concentrated on connectivity. This makes it easier to specify and analyse the system-level model. Variable `next` will always be a subgraph of `net` (6) and be an injective function (7) meaning a section can have at most one predecessor and at most one successor under `next`.

Each train may occupy several sections and this is modelled by the relation `occp` (8). Note that more than one train may be related to the same section by `occp` which is a potential hazard since trains occupying the same section may collide. Although the railway strives to avoid multiple occupancy of the same section, we feel that a pure model which absolutely prevented it would be unrealistic. The invariant does have a weaker safety constraint saying that if the front of `t1` is on the same section as the back of `t2`, then `t2` should be braking (13). This may be insufficient to avoid collisions since it doesn't take account of stopping distance, but it gives a flavour of a safety constraint for the purposes of this paper. The functions `front` and `back` map each train to the section occupied by its front and back respectively (9), (10). The function `braking` maps each train to a boolean representing whether or not it is braking.

Figure 2.5 describes some operations of the system-level model. The `Check` operation checks whether the brakes need to be applied to a train `t`. The brakes are applied if the front of `t` has no successor or its successor is already occupied[4]. The brakes are applied by setting `braking(t)` to `TRUE`. Regardless of whether the brakes need to be applied, the `Check` operation sets the `checked(t)` flag to true.

```
Check(t) =
    PRE t ∈ TRAIN
    THEN
      braking(t) := bool(front(t) ∉ dom(next) ∨ next(front(t)) ∈ ran(occp)) ||
      checked(t) := TRUE
    END;
EnterSection(t,s) =
    PRE t ∈ TRAIN ∧ s ∈ SECTION
    THEN
        SELECT checked(t) ∧ front(t) ∈ dom(next) ∧ s = next(front(t))
        THEN   front(t) := s || occp := occp ∪ {t ∑ s} || checked(t) := FALSE
        END
    END;
LeaveSection(t,s) = ⌡
SwitchChangeDiv(sw,s1,s2) =
    PRE sw, s1, s2 ∈ SECTION
    THEN
        SELECT sw ∈ switch ∧ sw ∉ ∈ ran(occp) ∧
               (sw ∑ s1) ∈ next ∧ (sw ∑ s2) ∈ net ∧ s1 ≠ s2
        THEN   next(sw) := s2
        END
    END;
SwitchChangeDiv(sw,s1,s2) = ⌡
```

Figure 2.5: Some operations of abstract railway.

[4] `t ∉ dom(next)` means that `next(t)` is undefined. `ran(occp)` is the set of sections that are occupied.

The `EnterSection` operation models the front of train `t` entering section `s`. Train `t` can enter section `s` provided `t` has been checked (since it last entered a section) and `t` has a successor section under `next`. The effect of the operation is to change the front section of the train to be `s`, add a mapping from `t` to `s` to the relation `occp` and reset the `checked(t)` flag.

Use of the `checked` flag means that the check operation must be applied to each train `t` at least once in between each `EnterSection` for `t`. In the real system, the train cannot be directly prevented from entering a new section and the only way of satisfying this ordering constraint is to ensure that the `Check` operation is performed before the train enters a new section. So this represents a hard real-time constraint.

The `SwitchChangeDiv` operation represents a switch change for a divergent switch and is modelled by the effect the switch change has on the `next` function. Parameter `sw` must be a switch which is currently connected to successor section `s1`. The effect of the operation is to update the `next` function so that `sw` is connected to `s2` instead. There is a further constraint on the occurrence of the switch change saying that the `sw` should not be currently occupied since changing an occupied switch has the potential to cause a derailment.

2.1.3.2 Modelling Time Properties

This section illustrates how to model time constraints in a specification [Butler02b], by using a simplified version of the case study presented in Section 1.2.1. The approach taken to modelling timing constraints is to include a clock variable representing the current time and an operation which advances this variable. The timing constraints are ensured to be satisfied by preventing the clock variable from progressing to a point at which the required properties would be violated. It would seem that one could always satisfy the properties by preventing time from progressing. However, one assumes that in the real system time cannot be prevented from progressing and thus it is an obligation on the final implementation to ensure that the timing properties are always satisfied in time. This is a fairly standard approach to dealing with time in formalisms. For example, this is similar to the approach taken by Abadi and Lamport [Abadi94]. Unlike Abadi and Lamport, this approach uses a discrete model of time rather than a continuous one. A discrete model is sufficient for our purposes since we are interested in ensuring that certain properties hold within fixed time bounds and since the control of the real system is based on fixed time cycles. When we say that something happens at a certain time `t`, what we mean is that it happened within time period `t`.

Consider the timing property on the emergency brakes again: *within `X` milliseconds of a section becoming restrictive, the emergency brakes of a train on that section are applied*. To represent this constraint in B, we include a variable which records the most recent time at which a section went restrictive, `rtime`. Then the requirement may be formalised in the following form (`t` is the current time):

```
Prop(t)  =  t ≥ rtime + X  ⇒  braking = true
```

The operation which progresses time is then guarded by the condition that the property holds in the new time period:

```
NextTime  =  SELECT  Prop(t+1)  THEN  t := t+1  END
```

Ultimately the real system will have to ensure that `Prop(t+1)` does indeed hold by the time `t+1` is reached.

```
MACHINE  RailwayT
SETS  SECTION, TRAIN, BLOCK
CONSTANTS  block, next, XX
PROPERTIES
   block ∈ SECTION → BLOCK ∧
   next ∈ BLOCK ⊙ BLOCK ∧
   XX ∈ NAT ∧ XX > 0
VARIABLES  occ, res, pos, cur, rtime, atime, barking
INVARIANT
   occ ⊆ SECTION ∧                  /* set of occupied sections */
   res ⊆ SECTION ∧                  /* set of restrictive sections */
   pos ∈ TRAIN → SECTION ∧          /* each train has a single position */
   cur ∈ NAT ∧                      /* The current time interval */
   rtime ∈ SECTION → (0..cur) ∧     /* time at which a section goes
                                       restrictive */
   atime ∈ TRAIN → (0..cur) ∧       /* time at which a train enters current
                                       position */
   braking ⊆ TRAIN ∧                /* set of braking trains */
   brakingProperty(cur)
```

Figure 2.6: RailwayT machine.

Figure 2.6 presents part of the system specification. The property `BrakingProperty(cur)` in the invariant, where `BrakingProperty(i)` is defined as follows:

```
BrakingProperty(i) =
      ∀tt.( tt ∈ TRAIN ∧ pos(tt) ∈ res ∧
            atime(tt) < rtime(pos(tt)) ∧ rtime(pos(tt)) + XX ≤ i
            ⇒ tt ∈ braking )
      ∧
      ∀tt.( tt ∈ TRAIN ∧ pos(tt) ∈ res ∧
            atime(tt) ≥ rtime(pos(tt)) ∧ atime(tt) + XX ≤ i
            ⇒ tt ∈ braking )
```

The braking property has two cases. If the train has arrived in the section before the section went restrictive, then the time bound is relative to the time at which the section went restrictive. If the train has arrived in the section after the section went restrictive, then the time bound is relative to the time at which the train arrived.

```
Arrive(t,s) =
   PRE t ∈ TRAIN ∧ s ∈ SECTION THEN pos(t) := s || atime(t) := cur END;

Restrict(s) = PRE s ∈ SECTION THEN res := res ∪ {s} END;

NextTime = SELECT BrakingProperty(cur+1) THEN cur := cur+1 END;

Brake(t) = PRE t ∈ TRAIN THEN braking := braking ∪ {t} END
```

Figure 2.7: RailwayT operations

Train arrival and sections going restrictive are modelled by the operations presented in Figure 2.7. Note that have simplified things considerably here. In our larger model, sections are

restrictive when certain other sections are occupied so that the `Arrival` operation causes relevant other sections to go restrictive. Also in the larger model, a train cannot arrive at an arbitrary section.

The operation which increases time is guarded by the braking property holding on the next time interval. In case the braking property would not hold in the next time interval, it can always be made to hold by some invocations of the operation which models application of the brakes.

2.2 Event B

The event B [Abrial98] approach is based on B notations of the B method and extends the methodological scope of initial concepts such as set-theoretical notations and generalised substitutions to take into account abstract systems. An abstract system is characterised by a (finite) list of variables possibly modified by a (finite) list of events; an invariant establishes properties satisfied by variables and maintained by the activation of events (with true guards) reacting to the environment. Abstract systems are close to action systems. A system model may be decomposed into sub-systems, representing both hardware and software components. At the lowest level, the model provides event based software level specification that can be used for a traditional software development in B or using any other language.

In event B, a model of a system has two parts:

- a *static* part – which describes the state of the system; and
- a *dynamic* part – which describes the events that may occur in the system.

Properties of the state are related to *safety*, whereas properties of events relate to *liveness*.

Besides its state, a model contains a number of events that show the way it may evolve. These events are only supposed to be observable; they are in no way actions that can be "called". Indeed, we are not describing the realisation here, all be it abstract, of a system. Rather, we are making a mathematical simulation, which allows us to reason about the future system we are going to construct. This reasoning is precisely what will allow us to analyse very early on the behaviour of our future system and to draw up a possible architecture.

Each event is composed of a guard and an action. The guard is the necessary condition under which the event may occur. In other words, once its guard holds, the occurrence of the event may be observed at any time (but it may also never be observed). As soon as the guard ceases to hold however, the event cannot be observed. The action, as its name indicates, determines the way in which the state variables will evolve when the event does occur.

It is possible for the guards of several events to hold simultaneously. From this point of view, the model presents a certain external non-determinism. Note that when several guards hold simultaneously, no two events cannot be observed to occur "together": events are atomic.

Finally, we must observe that the events are, *a priori*, asynchronous. Possible synchronisms are only the consequence of the actions of some on the guards of others.

Structure of an Event

An event is expressed in the following form:

```
e  =
    ANY x, y, … WHERE
        P(x, y, …, v, w, …)
    THEN
        S(x, y, …, v, w, …)
    END
```

For such an event, its guard corresponds to the existential predicate:

$$\exists\ (x, y, \ldots)\ .\ P(x, y, \ldots, v, w, \ldots)$$

In other words, the necessary (but insufficient) condition for the event *e* to take place with the current value of the state variables or constants `v, w, …` of the model, is that it be possible to assign to the local variables `x, y, …,` of the event `e`, some values making the predicate `P(x, y, …, v, w, …)` true. As can be seen *e* presents a certain latitude in the choice of possible values for the local variables `x, y, …` . We can speak here about internal nondeterminism.

The action `S(x, y, …, v, w, …)` is presented in the form of the simultaneous assignment of certain state variables `x, y, …` to certain expressions `E, F, …` depending upon the state of the system and the local variables `x, y, …` of the event. Note that any variables not mentioned in the list `x, y, …` do not change:

$$x, y, \ldots = E, F, \ldots$$

Sometimes, the event can have the simpler form:

```
xxx =
    SELECT
            P(v, w, … )
    THEN
            S(v, w, … )
    END
```

In this case, there are no variables local to the event, and the guard just corresponds to the condition `P(v, w, … )` holding on the state variables of the model.

Consistency: Preservation of the Invariant

Once a model has been built, it has to be proven to be consistent. Each event of the model has to be proven to preserve the invariant. More precisely, it must be proved that the action associated with each event modifies the state variables in such a way that the corresponding new invariant holds under the hypothesis of the former invariant and of the guard of the event. For a model with state variable `v` and invariant `I(v)`, and an event of the form:

```
ANY  x  WHERE
    P(x, v)
THEN
    v := E(x, v)
END
```

the statement to be proved is:

$$I(v) \wedge P(x, v) \Rightarrow I(E(x, v))$$

2.2.1 Example: Dealing with Events

The following example shows how to add/group/split events.

We consider here a simple system (M0). Its state is represented by a variable xx and one event may occur, evol, which led variable xx to evolve according to its definition domain, namely N.

```
SYSTEM M0
VARIABLES
   xx
INVARIANT
   xx ∈ NAT
INITIALISATION
   xx :∈ NAT
EVENTS
   evol = BEGIN xx ∈ (xx ∈ NAT) END
END
```

In the next refinement (M1), we express that our variable xx evolves according to the value of an other state variable, yy, which can be considered as a stimulus.

A new event is created, evoly, related to the evolution of the variable yy. In order to ensure that this new event may not take forever the control of the system and prevent event evol to occur, the number of occurrences of that event should be finite. We introduce in this case a finite set zz. Values for yy are chosen in this set and removed from it. When this set is empty, event *evoly* cannot be fired any more and variable yy keeps its last value. The VARIANT clause is introduced, containing an expression that should be decreased each time the new event evoly is fired and should reach zero in a finite number of steps. The VARIANT expression is the cardinal of the set zz.

Event evol is split into 3 events, each one describing the evolution of xx in different cases, depending on the value of yy. In fact, xx is modified in order to reach the value of yy.

The event refinement diagram is given below:

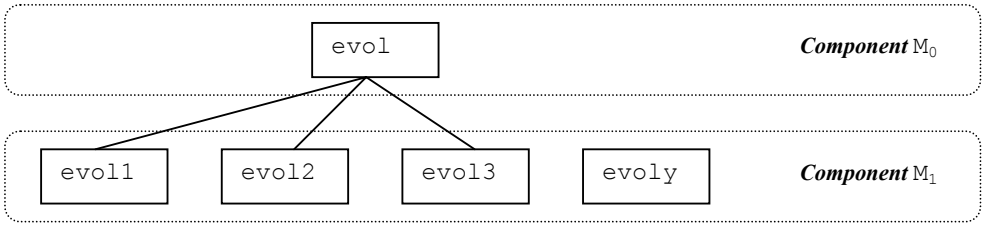


Figure 2.8: Events location in components M0 and M1.

In the B model (component M1), evol is explicitly replaced by evol1, evol2 and evol3. Each one is declared to refine evol. On the other hand, there is no need to express that evoly is new event since it can be deduced from previous component.

We obtain the event B model below.

```
REFINEMENT M1
REFINES M0
VARIABLES
   xx, yy, zz
INVARIANT
   xx ∈ NAT &
   yy ∈ NAT&
   zz ⊆ NAT &
   zz ∈ FIN(zz)
INITIALISATION
   xx :∈ NAT ||
   yy, zz ∈ (yy ∈ NAT  ∧  zz ⊆ NAT  ∧  zz ≠ ∅  ∧  yy ∈ zz  ∧  zz ∈ FIN(zz))
VARIANT
   card(zz)
EVENTS
   evol1 ref evol = SELECT xx=yy THEN skip END;
   evol2 ref evol = SELECT xx>yy THEN xx:=xx-1 END;
   evol3 ref evol = SELECT xx<yy THEN xx:=xx+1 END;
   evoly =
      SELECT zz ≠ ∅ THEN
         ANY val WHERE val ∈ zz THEN
            zz := zz - {val} ||
            yy := val
         END
      END
END
```

In the second refinement, evoly remains unchanged. evol1, evol2 and evol3 are grouped into one event to form an IF THEN ELSE substitution. This new event, evol4, is explicitly declared to refine evol1, evol2 and evol3.

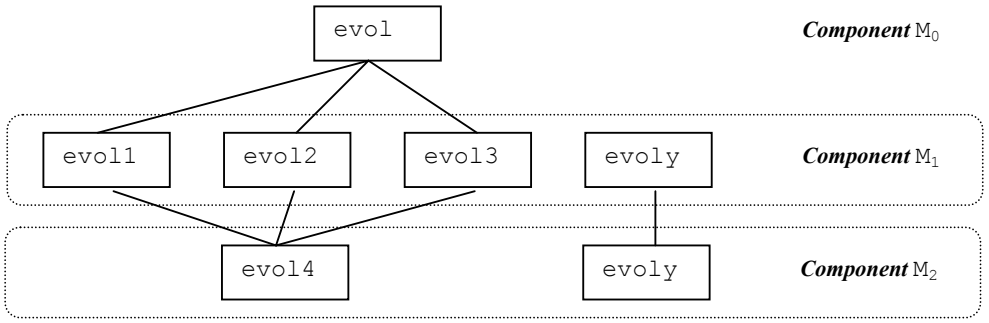


Figure 2.9: Events location in components M0, M1 and M2.

2.2.2 Example: Using Modalities

The following example shows how to use modalities.

We consider here a simple system (S0). Its state is represented by two variables xx and input. xx is linked with input such as, when input evolves (event ev1), then xx is likely to evolve too (events ev2 to ev7). We can consider that ev1 represents a kind a stimulus (from the environment) and ev2 to ev7 represent the response to this stimulus. Modalities can be used to write invariant properties restricted to some events (*i.e,*. ev2, ev3, … ev7). In this case, we would like to verify that the response complies with some general law, exhibited in the ESTABLISH clause.

```
SYSTEM
   S0
SETS
   EE = {e0, e1, e2, e3, e4};
   INPUTS = {v1, v2, v3}
VARIABLES
   xx,
   input
INVARIANT
   xx ∈ EE ∧
   input ∈ INPUTS
INITIALISATION
   xx := e0 ||
   input :∈ INPUTS
EVENTS
   ev1 =
      ANY in WHERE in ∈ INPUTS
      THEN
         input := in
      END
   ;
   ev2 = SELECT input=v1 ∧ xx = e0 THEN xx := e2 END;
   ev3 = SELECT input=v1 ∧ xx ≠ e0 THEN xx := e0 END;
   ev4 = SELECT input=v2 ∧ xx = e0 THEN xx := e2 END;
   ev5 = SELECT input=v2 ∧ xx ≠ e0 THEN xx := e3 END;
   ev6 = SELECT input=v3 ∧ xx = e1 THEN xx := e4 END;
   ev7 = SELECT input=v3 ∧ xx ≠ e1 THEN xx := e1 END
MODALITIES
   BEGIN
      ev2, ev3, ev4, ev5, ev6, ev7
   ESTABLISH
    input ∑ xx ∈ {
      (v1 ∑ e0), (v1 ∑ e2), (v1 ∑ e3),
      (v2 ∑ e2), (v2 ∑ e3),
      (v3 ∑ e1), (v3 ∑ e3), (v3 ∑ e4)
      }
   END

END
```

2.3 UML+B Notation

The U2B tool is a prototype tool to convert adapted forms of UML class diagrams and state chart diagrams into specifications in the B language. The aim is to use some of the features of UML diagrams to make the process of writing formal specifications easier, or at least more to the

average programmer. The translation relies on the precise expression of additional behavioural constraints in the specification of class diagram components and in state charts attached to the classes. These constraints are described in an adapted form of the B abstract machine notation. The type of class diagrams that can be converted is restricted in order to comply with constraints of the B-method without making the resultant B unnatural. The resulting UML model is a precise formal specification but in a form which is more friendly to the average programmer, particularly if they use the same UML notation for their program design work.

2.3.1 U2B Class Diagram Translator

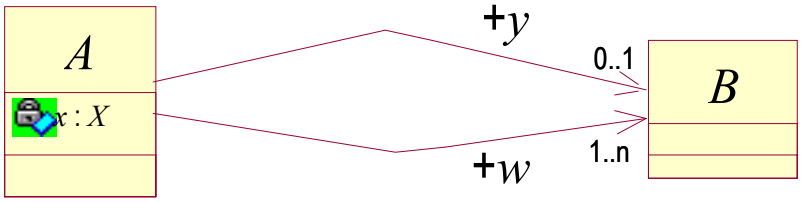
The U2B translator converts Rational Rose[5] UML Class diagrams [Rational00a], including attached state charts, into the B notation. U2B is a script file that runs within Rational Rose and converts the currently open model to B. It is written in the Rational Rose Scripting language, which is an extended version of the Summit Basic Script language [Rational00a, Rational00b]. U2B is configured as a menu option in Rose. U2B uses the object-oriented libraries of the Rose Extensibility Interface to extract information about the classes in the logical diagram of the currently open model. The object model representation of the UML diagram means that information is easily retrieved and the program structure can be based around the logical information in the class rather than a particular textual format. U2B uses Microsoft Word[6] to generate the B Machine files.

2.3.1.1 Translation of Structure and Static Properties

The translation of classes, attributes and operations is derived from proposals for converting OMT (Object Modelling Technique) to B [Meyer99]. However, since our aims are primarily to assist in the creation of a B specification rather than to generate a formal equivalent of a UML specification, our translation simplifies that proposed by [Meyer99]. This is achieved by restricting the translation to a suitable subset of UML models.

A separate machine is created for each class and this contains a set of all possible instances of the class and a variable that represents the subset of current instances of the class. Attributes and (unidirectional) associations are translated into variables whose type is defined as a function from the current instance to the attribute type (as defined in the class diagram) or associated class.

For example consider the following class diagram with classes *A* and *B*, where *A* has an attribute *x* and there is a unidirectional association from *A* to *B* with role *y* and 0 or 1 multiplicity at the target end. A second association, *w,* has a 'many' multiplicity:



[5]Rational Rose is a trademark of the Rational Software Corporation

[6]Microsoft Word97 is a trademark of the Microsoft Corporation.

This will result in the following machine representing all instances of *A*:

```
MACHINE    A
EXTENDS
        B
SETS
        ASET
VARIABLES
        Ainstances,
        x,
        w,
        y
INVARIANT
        Ainstances ⊆ ASET ⇸
        x ∈ Ainstances → X ⇸
        w ∈ Ainstances → POW1(Binstances) ⇸
        y ∈ Ainstances ⇓ Binstances
INITIALISATION
        Ainstances := ∅ ||
        x := ∅ ||
        w := ∅ ||
        y := ∅
```

Note that the multiplicity of the association *w* is handled as a function from instances of class *A* to sets of instances of class *B* using the power set operator (`POW` and `POW1`). The machine is initialised with no instances and hence all attribute and association functions are empty. A separate machine will be generated for class *B*.

Association multiplicities In UML, multiplicity ranges constrain associations. The multiplicities are equivalent to the usual mathematical categorisations of functions: partial, total, injective, surjective and their combinations. Note that the multiplicity at the target end of the association (class *B* in the example above) specifies the number of instances of *B* that instances of the source end, class *A,* can map to. This can be confusing when thinking in terms of functions because the constraint is at the opposite end of the association to the set it is constraining. The multiplicity of an association determines its modelling as shown in Table 2. We use functions to sets of the target class instances (*e.g.*, `POW(B)`) to avoid non-functions. Note that `0..n` is assumed unless otherwise specified in UML.

Association Representations in B for Different Multiplicities		
Ai *and* ***Bi*** *are the current instances sets of class A and B respectively (i.e., Ainstances and Binstances) and* ***f*** *is a function representing the association (i.e., the role name of the association with respect to the source class, A).*		
UML association multiplicity	**Informal description of B representation**	**B invariant**
`0..n` → `0..1`	partial function to *Bi*	$Ai \Downarrow Bi$
`0..n` → `1..1`	total function to *Bi*	$Ai \rightarrow Bi$
`0..n` → `0..n`	total function to subsets of *Bi*	$Ai \rightarrow \mathtt{POW}(Bi)$
`0..n` → `1..n`	total function to non-empty subsets of *Bi*	$Ai \rightarrow \mathtt{POW1}(Bi)$
`0..1` → `0..1`	partial injection to *Bi*	$Ai \rightarrowtail Bi$
`0..1` → `1..1`	total injection to *Bi*	$Ai \odot Bi$
`0..1` → `0..n`	total function to subsets of *Bi* which do not intersect	$Ai \rightarrow \mathtt{POW}(Bi) \twoheadrightarrow \mathtt{inter}(\mathtt{ran}(f)) = \varnothing$
`0..1` → `1..n`	total function to non-empty subsets of *Bi* which do not intersect	$Ai \rightarrow \mathtt{POW1}(Bi) \twoheadrightarrow \mathtt{inter}(\mathtt{ran}(f)) = \varnothing$
`1..n` → `0..1`	partial surjection to *Bi*	$Ai \; ⤀ \; Bi$
`1..n` → `1..1`	total surjection to *Bi*	$Ai \; ⤀ \; Bi$
`1..n` → `0..n`	total function to subsets of *Bi* which cover *Bi*	$Ai \rightarrow \mathtt{POW}(Bi) \twoheadrightarrow \mathtt{union}(\mathtt{ran}(f)) = Bi$
`1..n` → `1..n`	total function to non-empty subsets of *Bi* which cover *Bi*	$Ai \rightarrow \mathtt{POW1}(Bi) \twoheadrightarrow \mathtt{union}(\mathtt{ran}(f)) = Bi$
`1..1` → `0..1`	partial bijection to *Bi*	$Ai \; ⤖ \; Bi$
`1..1` → `1..1`	total bijection to *Bi*	$Ai \; ⤖ \; Bi$
`1..1` → `0..n`	total function to subsets of *Bi* which cover *Bi* without intersecting	$Ai \rightarrow \mathtt{POW}(Bi) \twoheadrightarrow \mathtt{union}(\mathtt{ran}(f)) = Bi \twoheadrightarrow \mathtt{inter}(\mathtt{ran}(f)) = \varnothing$
`1..1` → `1..n`	total function to non-empty subsets of *Bi* which cover *Bi* without intersecting	$Ai \rightarrow \mathtt{POW1}(Bi) \twoheadrightarrow \mathtt{union}(\mathtt{ran}(f)) = Bi \twoheadrightarrow \mathtt{inter}(\mathtt{ran}(f)) = \varnothing$

Table 2: How associations are represented in B for each possible multiplicity constraint.

In Figure 2.10 a mapping represents an association between the classes *A* and *B* with multiplicity `0..n → 0..1`. The representation in the B notation is a partial function. It is not a total function because a_4 doesn't map to anything in B (as indicated by the 0 at the right hand end of `0..n → 0..1`). It is not injective because b_2 is mapped to by both a_2 and a_3 (as indicated by the n at the left hand end of `0..n → 0..1`). It is not surjective because b_3 is not mapped to by anything in *A* (as indicated by the `0` at the left hand end of `0..n → 0..1`).

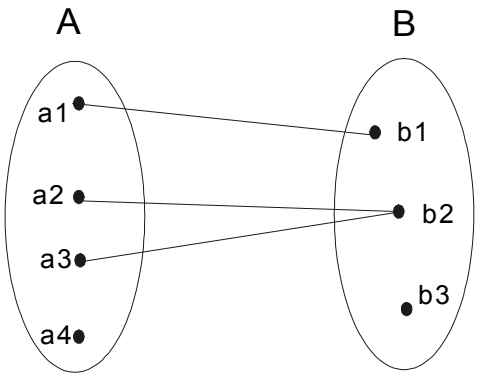
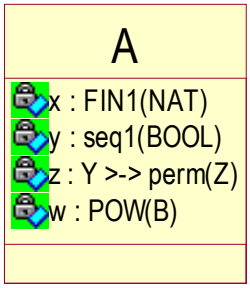
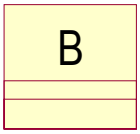


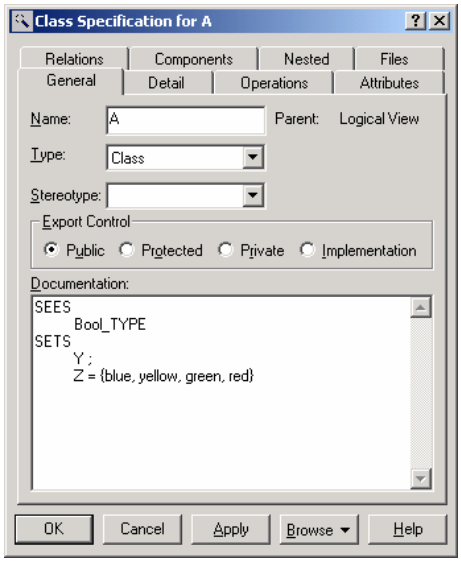
Figure 2.10: Mapping representing a `0..n → 0..1` association.

Attribute types Attribute types may be any valid B expression that defines a set. This includes predefined types such as `NAT`, `NAT1`, `boolean` and `string` (if translating to B-Core B, the appropriate B library machines must be referenced via a `SEES` clause in the class's specification documentation window) functions, sequences, power sets, instances of another class (referenced by the class name) or enumerated or deferred sets defined in the class specification documentation window. If the type involves another class (and there is no unidirectional path of associations to that class) the machine for that class will be referenced in a `USES` clause so that its current instances set can be read. If there is a path of unidirectional associations to the class it will be extended (`EXTENDS`) by this machine in order to represent the association and this will provide access to the instances set. (Note that only unidirectional associations are interpreted as associations. Unspecified or bi-directional associations are ignored and can therefore be used to indicate type dependencies diagrammatically if required). Any references to the class in type definitions of variables or operation arguments will be changed to the current instances set for that class.

For example, the following shows a class that has an attribute `x` of type, non-empty finite subset of natural numbers. It has an attribute `y` that is of type, non-empty sequence of booleans. The library machine `Bool_TYPE` has been referenced via a `SEES` clause in the class's documentation window (this would not be necessary for Atelier-B). It has an attribute `z` that has type, total injection from `Y` to permutations of `Z`. A `SETS` clause has been added to the class's documentation window that defines `Y` as a deferred set and `Z` as an enumerated set.



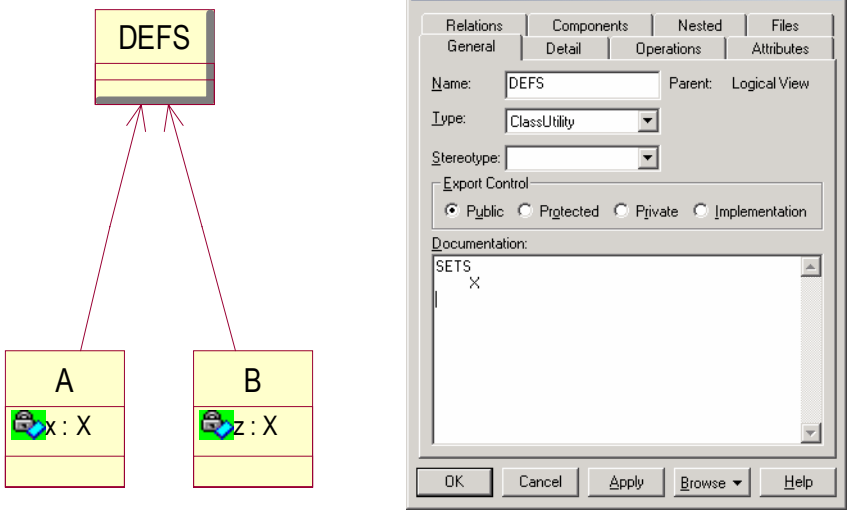




Note that `Export Control` settings in the class specification are not used in the U2B translation. The corresponding B machine for class *A* is shown below.

```
MACHINE   A
SEES
       Bool_TYPE
USES
       B
SETS
       ASET ;
       Y ;
       Z = {blue, yellow, green, red}
VARIABLES
       Ainstances,
       x,
       y,
       z,
       w
INVARIANT
       Ainstances ⊆ ASET ⇸
       x ∈ Ainstances → FIN1(NAT) ⇸
       y ∈ Ainstances → seq1(BOOL) ⇸
       z ∈ Ainstances → Y ⊙ perm(Z) ⇸
       w ∈ Ainstances → POW(Binstances)
```

Global Definitions It is often useful to define types as enumerated or deferred sets for use in many machines. We use *class utilities* for this. In UML, a class utility is a class that has no instances, only static (class-wide) operations and attributes. The U2B translator creates a machine for each class utility and copies any text in the specification documentation window of its class specification into the machine. Hence definitions, sets and constants can be described in B clauses in the documentation window. Any machines that reference items defined in this way must have an association to the class utility. This association will not be interpreted as an association to an ordinary class). In the following example a class utility `DEFS` is used to define a set `X` that is used as a type by two other classes.



The corresponding machine for class utility `DEFS` is:

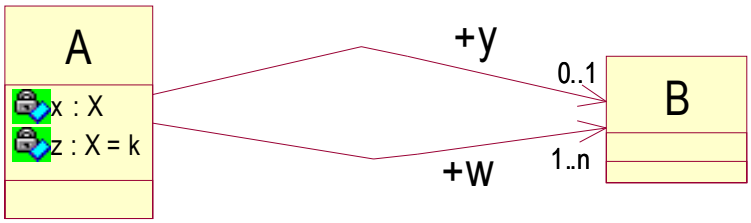
```
MACHINE    DEFS
SETS
          X
END
```

The machines for classes *A* and *B* will reference `DEFS` via a `SEES` clause:

```
SEES
        DEFS
```

Local Definitions Sets can also be defined locally to a class in the class's specification documentation window. In fact, any valid B clause can be added in this window. For example, we use this method to specify invariants for the class. Each clause must be headed by its B clause name in capitals and starting at the beginning of a line, the text that follows that clause, up until the next clause title (if any) will be added to the appropriate clause in the machine. Any text before the first clause is treated as comment and added as such at the top of the machine

Instance Creation and Initialisation of Attributes and Associations. A create operation is automatically provided for each class machine so that new instances can be created. This picks any instance that is not already in use, adds it to the current instances set, and adds a maplet to each of the attribute/association relations mapping the new instance to the appropriate initial value. Note that, according to our definition (via translation) of class diagrams, association means that the source class is able to invoke the methods of the target class. The example below is similar to the first example but class *A* has an additional attribute *z*, that has an initial value *k*.




```
Return ← Acreate =
    PRE
        Ainstances ≠ ASET
    THEN
        ANY new
        WHERE
            new ∈ ASET - Ainstances
        THEN
            Ainstances := Ainstances ∪ {new} ||
            x(new):∈ X ||
            z(new):= k ||
            w(new):∈ POW1(Binstances) ||
            Return := new
        END
    END
```

Note that, because `x` has no initial value specified, it is initialised non-deterministically to any value of the type `X` (`:∈` means any value belonging to). Similarly the association *w* must be initialised to a non empty set because its multiplicity may be greater than one but is definitely greater than zero. (Currently, we have no means of specifying initial values for associations). It is initialised non-deterministically to any non-empty subset of B instances. The association *y* is not initialised because its multiplicity is 0 or 1 and it may, therefore, be undefined. Initially the new instance will have no maplet representing the association *y*.

Singular Classes Often, a B machine models a single generic instance of an entity, rather than an explicit set of instances (in the same way that a class in UML leaves instance referencing implicit). The resulting specification is simpler and clearer for not modelling multiple instances. The U2B translator creates a single-instance machine if the class multiplicity (cardinality) is set to `1..1` in the UML class specification. Note that this can only be done at the top level of a structure since at lower levels the instance set is used for referencing by the higher level. Below is shown the machine representing class *A* from the first example above if the class's multiplicity is set to `1..1`. Note that there is no modelling of instances, the type of attributes is simpler because it is no longer necessary to map from instances to the attribute type. There is no instance creation operation; attributes are initialised in the machine initialisation clause.

```
MACHINE   A
EXTENDS
       B
VARIABLES
       x,
       w,
       y
INVARIANT
       x ∈ X ∧
       w ∈ POW1(Binstances) ∧
       y ∈ Binstances
INITIALISATION
       X :∈ X ||
       w :∈ POW1(Binstances) ||
       y := ∅
END
```

Restrictions The B method imposes some restrictions on the way machines can be composed. These restrictions ensure compositionality of proof. Their impact is that no write sharing is allowed at machine level (*i.e.,* a machine may only be included or extended by one other machine). Also, the inclusion mechanism of B is hierarchical so that, if `M1` includes `M2`, then `M2` cannot, directly or transitively, include `M1`. We reflect these restrictions in the UML form of the specification, which must therefore be tree like in terms of unidirectionally related classes. Non-navigable (and bi-directional) associations are ignored but may be used to illustrate the use of another class as a type (*i.e.,* read access only). However, multiple, parallel associations between the same pair of classes are permitted.

Although we would like to adhere to the UML class diagram rules as much as possible, since our aim is to make B specification more approachable rather than to formalise the UML we are relatively happy to impose restrictions on the diagrams that can be drawn. That is, we only define translations for a subset of UML class diagrams. Other authors [Facon96, Meyer99, Meyer00, Nagui94, Shore96] have suggested ways of dealing with the translation of more general forms of class diagrams. However, the structures of B machines that result from these more general translations can be cumbersome. If the specification were written directly in B, it would be highly unlikely that the resulting B would have this form. Since we also desire a usable B specification we prefer to restrict the types of diagrams that can be drawn.

2.3.1.2 Dynamic Behaviour

The dynamic behaviour modelled on a class diagram that is converted to B by U2B is embodied in the behaviour specification of classes operations and in invariants specified for the classes. UML does not impose any particular notation for these operation and invariant constraint definitions; they could be described in natural language or using UML's Object Constraint Language (OCL). However since we wish to end up with a B specification it makes sense to use bits of B notation to specify these constraints. The constraints are specified in a notation that is close to B notation but has to observe a few conventions in order for it to become valid B within the context of the machine produced by U2B. When writing these portions of B the writer should not need to consider how the translation would represent the features (associations, attributes and operations) of the classes. Also we felt we should follow the more object-oriented conventions of implicit self-referencing and use of the dot notation for explicit instance references. Therefore,

when writing the constraints, a dot notation is used to reference the ownership of features. This is illustrated in examples below.

Invariant Unfortunately there is no dedicated text box for a class invariant in Rational Rose. One suggestion is to put invariant constraints in a note attached to the class [Warmer99], but notes are treated as an annotation on a particular view in Rational Rose and not part of the model. This makes them difficult to access from the translation program and unreliable should we extend the conversion to look at other views. Therefore we include the invariants as a clause in the documentation text box of the class' specification window. The invariants are generally of two kinds, *instance invariants* (describing properties that hold between the attributes and relationships within a single instance) and *class invariants* (describing properties that hold between different instances). To deal with instance invariants, and keeping with the implicit self-reference style of UML, we chose to allow the explicit reference to this instance to be omitted. U2B will add the universal quantification over all instances of the class automatically. For class invariants, the quantification over instances is an integral part of the property and must be given explicitly. Hence, U2B will not need to add instance references.

For example, if `bx ∈ NAT` is an attribute of class *B* then the following invariant could be defined in the documentation box for class *B*:

```
bx < 100 ∧
∀(b1,b2).((b1 ∈ B ∧ b2 ∈ B ∧ b1 ≠ b2) => (b1.bx ≠ b2.bx)
```

This would be translated to:

```
∀(thisB).(thisB ∈ Binstances  ⇒
   bx(thisB) < 100 ∧
   ∀(b1,b2).((b1 ∈ Binstances ∧ b2 ∈ Binstances ∧ b1 ≠ b2)
⇒ (bx(b1) ≠ bx(b2))
)
```

The translation has added a universal quantification, `thisB`, over all instances of *B* and this is used in the first part of the invariant. It is not used in the second part where the invariant already references instances of class *B*. (Note that currently the translator adds one universal quantification for the entire invariant whether or not it is needed).

Operation Semantics Operation preconditions are specified in a textual format attached to the operation within the class. Details of operation behaviour are specified either in a textual format attached to the operation, or in a state chart attached to the class. Operation behaviour may be specified completely by textual annotation, completely by state chart transitions, or by a combination of both composed as simultaneous specification.

Operation Textual behaviour specification In Rational Rose, *Specifications* are provided for operations (as well as many other elements) and these provide text boxes dedicated to writing pre-conditions and semantics for the operation. Although Rational Rose also provides a *post-*

condition text box , this is not currently used, as the *semantics* box suit the pseudo-operational style of B better.

Operations need to know which instance of the class they are to work on. This is implicit in the class diagram. The translation adds a parameter *thisCLASS* of type *CLASSinstances* to each operation. This is used as the instance parameter in each reference to an attribute or association of the class.

In the above example, *set_y* might have the following precondition:

```
i > y.bx
```

and semantics

```
y.b_op(i) ||
IF y.bx <100
THEN
    out := FALSE
ELSE
    out := TRUE
END
```

which would be translated to

```
i > bx(y(thisA))
```

and

```
b_op(y(thisA)) ||
IF bx(y(thisA)) <100
THEN
    out := FALSE
ELSE
    out := TRUE
END
```

Operation Return Type UML operation signatures contain a provision for specifying the type for a value returned by the operation. Since B infers this from the body of the operation we use it instead to name the identifiers that represent operation return values. The string entered in the return type field for the operation will be used as the operation return signature in the B machine representing the class. For example, the *set_y* operation in the above class diagram has its return field set to *out*. The operation signature for `set_y` in the B machine `A` will be :

```
out <--  set_y (thisA,i) =
```

State chart Behavioural Specification For classes that have a strong concept of state change, a state chart representation of behaviour is appropriate. In UML a state chart model can be attached to a class to describe its behaviour. A state chart model consists of a set of states and transitions that represent the state changes that are allowed. If a state chart model is attached to a class the U2B translator combines the behaviour it describes with any operation semantics described in the operation specification semantics windows. Hence operation behaviour can be defined either in the *operation semantics* window or in a state chart model for the class or in a combination of both.

The name of the state chart model is used to define a state variable. (Note that this is not the name of a state chart diagram, several diagrams could be used to draw the state chart model of a class). The collection of states in the state chart model is used to define an enumerated set that is used in the type invariant of the state variable. The state variable is equivalent to an attribute of the class and may be referenced elsewhere in the class and by other classes. State chart transitions define which operation call causes the state variable to change from the source state to the target state, *i.e.*, an operation is only allowed when the state variable equals a state from which there is a transition associated with that operation. To associate a transition with an operation, the transition's name must be given the same name as the operation. Additional guard conditions can be attached to a transition to further constrain when it can take place. All transitions cause the implicit action of changing the state variable from the source state to the target state. (The source and target state may be the same). Additional actions (defined in B) can also be attached to transitions. The translator finds all transitions associated with an operation and compiles a `SELECT` substitution of the following form:

```
SELECT statevar=sourcestate1 ∧ sourcestate1_guards
THEN statevar:=targetstate1 || targetstate1_actions
WHEN statevar=sourcestate2 ∧ sourcestate2_guards
THEN statevar:=targetstate2 || targetstate2_actions
<etc>
END ||
```

This is composed with the operation pre-condition and body (if any) from the textual specification in the operation's `pre-condition` and `semantics` windows:

Let `Popw` be the precondition in the operation `precondition` window, `Sosw` be the operation body from the `operation semantics` window and `Gstc` the `SELECT` substitution for this operation composed from the state chart. Then the translator will produce the following operation:

```
PRE
   Popw
THEN
   Gstc ||
   Sosw
END
```

This can be represented more succinctly in B as:

```
Popw  |  (Gstc || Sosw)
```

Hence the precondition, `Popw`, has precedence and, if false, the operation will abort. If an event B style systems simulation is desired, the specifier should take care not to define pre-conditions that conflict with the transition guards. (For example, if an event only occurs if an attribute `bx` is

positive, and this is modelled by a guarded transition; adding the pre-condition `bx > 0` would change the meaning of the model to represent a system where if the event occurs the operation aborts).

Note that it would be entirely valid (although somewhat obtuse) to write a pre-condition within the *operation semantics* window: `Sosw = Posw | S1osw`. However, preconditions take precedence in simultaneous substitutions, so

```
(Gstc || (Posw | S1osw))  =  Posw | (Gstc || S1osw).
```

Hence, writing the precondition in the `operation semantics` window is equivalent to writing it in the `precondition` window. It has the same precedence and possible conflicts with the operation guards derived from the state chart. We feel that writing the precondition in the `operation semantics` window should be discouraged because the precedence may not be obvious to readers of the specification.

If the pre-condition (`Popw ∧ Posw`) is true, then the guard from `Gstc` takes precedence over the simultaneous substitution `Sosw`. This means that the textual operation body from the operation semantics window, although defined separately from the state chart and not associated with any particular state transition, is only enabled when at least one of the state transitions is enabled. That is, if

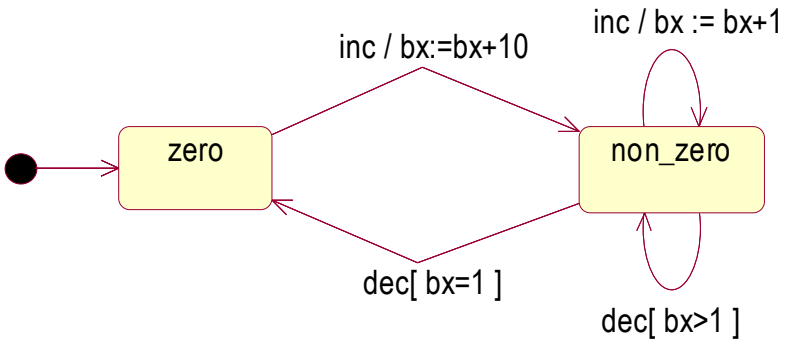
`Gstc` = (G1 ⇒ T1) □ .... □ (Gn ⇒ Tn)

then,

`(Gstc || Sosw) = ( G1 ⇒ ( T1 || Sosw ) ) □ … □ ( G ⇒ ( Tn || Sosw ) )`

where □ represents choice.

Actions should be specified on state transitions when the action is specific to that state transition. Where the action is the same for all that operation's state transitions, it may be specified in the operation semantics window in order to avoid repetition. The chart below illustrates how a state chart can be used to guard operations and define their actions. It also shows how common actions can be defined in the operation semantics window and how a pre-condition could upset the constraints imposed by the state chart.



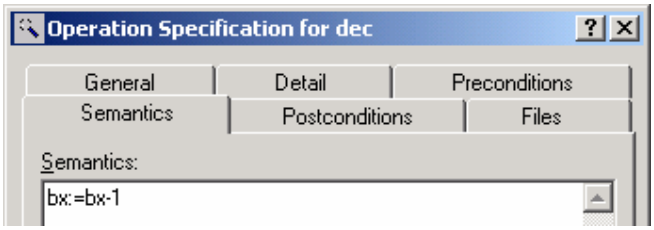
The state chart has two states, *zero* and *non_zero*. The implicit state variable, `b_state` (the name of the state chart model) is treated like an attribute of type `B_STATE = {zero, non_zero}`. An invariant defines the correspondence between the value of the attribute *bx* and the state *zero*. When an instance is created its `b_state` is initialised to `zero` because there is a transition from an initial state to *zero*.

```
MACHINE   B
SETS
       BSET;
       B_STATE={zero,non_zero}
VARIABLES
       Binstances,
       b_state,
       bx
INVARIANT
       Binstances ⊆ BSET ∧
       b_state ∈ Binstances → B_STATE ∧
       bx ∈ Binstances → NAT ∧
       ∀(thisB).(thisB ∈ Binstances ⇒
        (b_state(thisB)=zero) ⇔ (bx(thisB)=0)
        )
INITIALISATION
       Binstances := ∅ ||
       b_state := ∅ ||
       bx := ∅
OPERATIONS
Return  ← Bcreate =
   PRE
       Binstances ≠ BSET
   THEN
       ANY new
       WHERE
          new ∈ BSET - Binstances
       THEN
          Binstances := Binstances ∪ {new } ||
          b_state(new):=zero ||
          bx(new) :∈ NAT ||
          Return := new
       END
   END
;
```

Operation `inc` can occur in either state. Its action is different depending on the starting state and so actions have been defined on the transitions and are combined with the change state action.

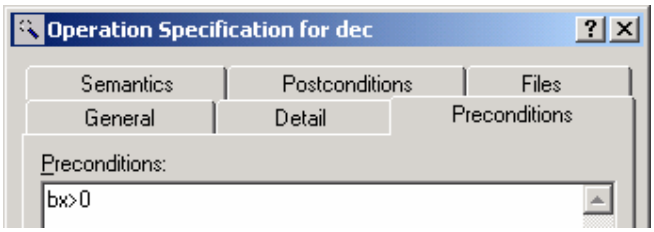
```
inc (thisB) =
   PRE
       thisB ∈ Binstances
   THEN
       SELECT b_state(thisB)=zero
       THEN   b_state(thisB) := non_zero ||
              bx(thisB):=bx(thisB)+10
       WHEN   b_state(thisB)=non_zero
       THEN   bx(thisB) := bx(thisB)+1
       END
   END
```

Operation `dec` has two guarded alternatives when in state `non_zero` but does not occur while in state `zero`. Since the action is the same for both transitions it has been defined in the *operation semantics* window.




```
dec (thisB) =
    PRE
        thisB ∈ Binstances
    THEN
        SELECT b_state(thisB)=non_zero ∧
               bx(thisB)=1
        THEN   b_state(thisB):=zero
        WHEN   b_state(thisB)=non_zero ∧
               bx(thisB)>1
        THEN   skip
        END ||
        bx(thisB):=bx(thisB)-1
    END
END
```

If we had put the pre-condition $bx > 0$ in the *operation specification precondition* window (or even in the *operation semantics* window), the guard would no longer function since the precondition would fail resulting in an abort when $bx = 0$.




```
dec (thisB) =
    PRE
        thisB ∈ Binstances ∧
        bx(thisB)>0
    THEN
        SELECT b_state(thisB)=non_zero ∧
               bx(thisB)=1
        THEN   b_state(thisB):=zero
        WHEN   b_state(thisB)=non_zero ∧
               bx(thisB)>1
        THEN   skip
        END ||
        bx(thisB):=bx(thisB)-1
    END
```

This could be avoided by repeating the precondition and decrement substitution in the action field of each *dec* transition on the state chart in which case the guard would take precedence.

3 Constructing Models

In this chapter we start by presenting in more detail each of the case studies developed within the MATISSE project. We divide the rest of the chapter into two different types of models: system-level models and computation.

3.1 System-Level Models

A system-level model starts with an abstract formal model of the overall system to be controlled, including relevant control behaviour. Then this high-level model is refined, adding detail about interactions between sub-components of the system, so that formal models may be decomposed into the system under control and the embedded controllers. Thus, rather than starting with formal models of the embedded controllers and verifying that they interact with the system under control, it starts with a model of the desirable behaviour of the overall system and use refinement to derive models of the embedded controllers. The refinement is typically performed in several steps, and at each step one has to verify that any behaviour of the refined model is allowed by the previous model, thus ensuring that the final detailed model is correct with respect to the original system-level model. The advantage of this approach is that is makes it easier to deal with the complexity of the designs. Starting with a simple system-level model and introducing extra complexity one-step at a time. The system-level model is particularly suited to systems that consist of multiple distributed controllers. The system level model abstracts away from any distribution, and only introduces it in refinement steps.

3.1.1 Translating Informal Properties into B

This section presents the properties for the first three levels of abstraction for the railway case study; in total there are eleven levels.

The railway system is described using several level of abstraction: the first level (the highest) being the most abstract one and subsequent levels provide increasing detail.

Each property is associated with a number *Pi-j* which means it is the *j*th property of the *i*th level of abstraction. Using the same numbering system, *SP* denotes safety property; *A*, automaton; and *E*, event.

There is no real difference between safety properties and “normal” properties, since safety properties make sense thanks to “normal” properties. In a way, all properties are “safety properties”, but a distinction is made in order to highlight the most significant properties related to safety needs.

First Level of Abstraction

At this level, the system just contains trains, a track, switches and floodgates. Safety is not ensured at this level, since *collision and derailment might occur*. There are three properties:

P1-1: Trains are running on the track.

P1-2: Switches are located on the track.

P1-3: Floodgates are located on the track.

Second Level of Abstraction

At this level, trains have several driving modes, switches and floodgates can operate, special areas are named (shunt zone, portion of track,…). Derailments are avoided but *collisions are still possible* (with a train or a floodgate).

P2-1: The track is divided into track circuits, called CDV (Circuit De Voie, in French). There is no intersection between two track circuits. The union of all track circuits is the entire track, this means there is no gap between two track circuits.

P2-2: Trains have a number of control modes, namely FCM (Full Control Mode), LCM (Limited Control Mode) and SCM (Stop Control Mode). They are each decomposed according to different driving modes, namely

- AM (Automatic Mode), DTO (Driverless Turn-round Operation) and CM (Coded Manual mode) for FCM;
- RMF (Restricted Manual Forward mode) and RMR (Restricted Manual Reverse mode) for LCM; and
- Stand By and Shut Down for SCM.

<u>If no crossed points are present on the track:</u>

P2-3: Switches are located on special track circuits, called Shunt Zones. A Shunt Zone is connected to three track circuits. A track circuit without switches is connected to only two track circuits.

<u>If crossed points are present on the track:</u>

P2-3 (bis): Switches are located on special track circuits, called Shunt Zones. A Shunt Zone is connected with three or four track circuits (for crossed point zones). A track circuit without switches is connected to only two track circuits.

SP2-1: Derailment of a train must be avoided. Thus, the position of a switch cannot change if there is a train (running or stopped) on the Shunt Zone of the switch.

E2-1: The state of a switch can change (two locked positions, an unknown position and an intermediate position), in accordance with SP2-1.

3.1.2 UML Development Incorporating Safety Aspects

The starting point of a system development is an informal specification describing the services desirable from the eventual product, with no or few hints about the various software and hardware parts. We consider this to be the most general starting point for the development process of a control system. The most important role is played by UML [UML1.4] in the first phase of the process, when a system specification is determined from the initial requirements. We define a method for using the UML diagrams for designing a control system specification

[Petre00]. An essential contribution at this level is the embedding of the safety analysis within the UML method.

We depict the informal requirements of the healthcare case study with UML diagrams. The functional requirements of the system are captured together with their relationships in a use case diagram. The reliability and safety issues are given in the specification of the use cases as structured English text. Then the logically related use cases are determined, and grouped together into a control system component in component and class diagrams. The behaviour of the component is specified with a statechart diagram. The informal specification of the eventual system is thus gradually captured and made more and more precise throughout these diagrams.

3.1.2.1 Functional Requirements Capturing Safety Issues

The *functional requirements* of the system are depicted, together with their relationships in a use case diagram as presented in Figure 3.1. Each use case expresses a service that the system will eventually provide to a user. In Figure 3.1 we have a system with three use cases.

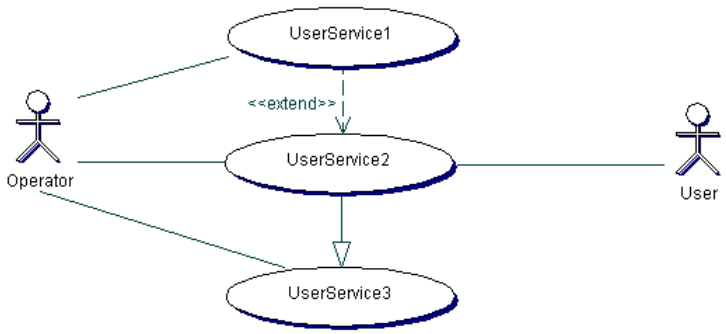


Figure 3.1: A use case diagram.

To exemplify the integrated methodology we create the controlling software for an analysing device, here called the Analyser. The Analyser is restricted to move vertically and to analyse blanks/plates. The simplified operations form the protocol of the analysing device. This protocol is described below. A typical protocol is:

1) Move to middle position.

2) Receive plate.

3) Move to upper end position.

4) Analyse the plate.

5) Move to lower end position.

6) Deliver plate.

7) Repeat from Step 1.

Hence, the functional requirements of the Analyser are to analyse plates, to receive plates from and deliver plates to the Robot and to move the operating table to its three possible positions. A use case diagram of the Analyser is given in Figure 3.2.

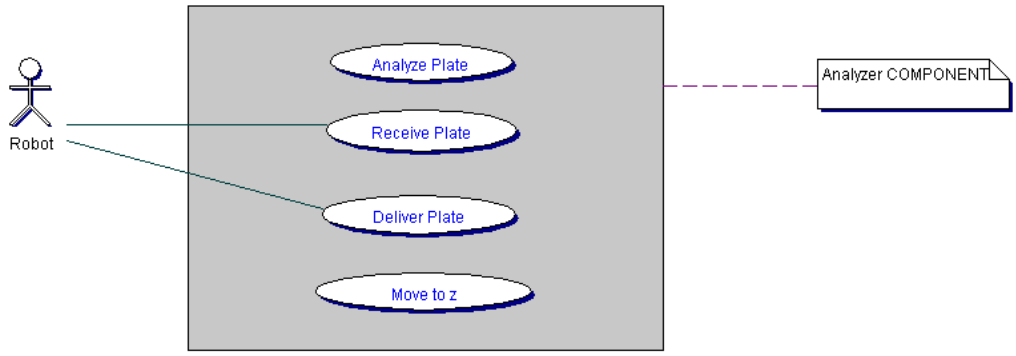


Figure 3.2: Use case diagram for the Analyser.

The Robot and the Analyser interact by the Robot loading a blank to the Analyser *receiving* it, as well as by the Robot unloading a blank from the Analyser *delivering* this blank. The Robot is the active partner while the Analyser is the passive partner during this interaction.

Besides these functional services we also capture, at a reasonable level of detail, the *reliability and safety issues* related to the determined functional services. These issues are not captured in the diagram, but in the specification of the use cases, most commonly expressed as structured English text. Hence, at this level the vocabulary of the whole system is fixed, an overview statement is defined, and the goals, attributes and basic functions, as well as their possible deviations are all captured by careful systematic analysis. The Analyser in our example is seen as a small part of the whole system.

Goals

We consider the following goals for the operator and user (as in Figure 3.1) of the whole system in our case study.

1) Easy operation of the whole assembly by humans.

2) Fast and accurate control of the mechanical parts.

3) Fast and accurate analysis of the protocol results.

4) Automatic inventory of the processing operations (perhaps stored in a log).

5) Careful functional and safety analysis for simulating the protocol of the Analyser operation.

The first four goals are common to the whole project and refer to the eventual software. The fifth goal refers essentially to the process of producing the eventual software. We apply a combination of formal methods (the B Method and B-action systems), UML and safety analysis for checking the quality of the eventual software and for producing this software, which we could then claim to be correct. This fifth goal is the most important to us.

Basic Functions

The basic functions for the eventual software of the healthcare case study are as follows:

1) Capture the protocol information when entered.

2) Communicate the protocol information with the mechanical parts.

3) Record the running protocol.

4) Inform about failures and the failure points.

These are the functions we have identified for the controlling software of the system. In this case study we are only interested in functions 1, 2 and 4.

We assume that the position detection of the operating table is absolutely reliable in the Analyser. The *risks* and *dependencies* of the system are listed during the development process.

Use Cases Incorporating Possible Failures

We have identified four use cases, *move*, *receive*, *analyse* as well as *deliver* for the Analyser as shown in Figure 3.2. We consider these use cases as subroutines, since it is more convenient with subroutines for the purpose of safety and reliability analyses.

Subroutine *Move* captures the moving of the operating table to the position *pos*.

Typical course of events:

1. Analyser receives the command to *move to position pos* from protocol.
2. System checks if pos is a valid position; if not, then MF1.
3. System checks if it is safe to move; if not then MF2.
4. System checks if current_pos does not equal pos and further moves are allowed; if not then MF3.
5. Analyser moves to position pos.
6. System checks if current_pos = pos; if not then MF4.
7. Signal move_ok.

Failure reports:

- MF1: Input parameter pos is outside the valid range.
  - Remedy: User changes parameter pos, identifies the cause and resumes or aborts calling protocol execution.
- MF2: Robot is not safe.
  - Remedy: Wait a while for the Robot to become safe.
- MF3: Erroneous move command. Protocol failure.
  - Remedy: Check and change protocol or abort.
- MF4: Analyser has not reached position pos.
  - Remedy: Calibrate and resume or abort calling protocol execution

Subroutine *Receive* captures the receiving of a plate from the Robot, provided the operating table is at its middle position and there is no plate in the Analyser.

Typical course of events:

1. Analyser receives command *receive* from protocol.
2. System checks if Robot has a plate available; if not then RF1.
3. System checks if current_pos = middle position to Receive; if not RF2.
4. System checks if the Analyser is empty; if not then RF3.
5. Robot loads the Analyser.
6. System checks if plate has been received; if not then RF4.
7. Signal receive_ok.

Failure reports:

- RF1: Robot is not ready to load the Analyser.
  - Remedy: Wait a while for the robot to become ready.
- RF2: Current_pos is not the middle position to Receive.
  - Remedy: User identifies the cause and moves the Analyser manually.
- RF3. Plate is present.
  - Remedy: User identifies the cause, and unloads the Analyser manually or aborts.
- RF4. The Analyser has not been loaded.
  - Remedy: User identifies the cause, loads the Analyser manually and resumes or aborts.

Subroutine *Analyse* captures the analysing of a plate, provided the operating table is at its upper end position and there is a plate in the Analyser.

Typical course of events:

1. Analyser receives command *analyse* from protocol.
2. System checks if a plate is present in the Analyser; if not then AF1.
3. System checks if current_pos = upper end position to Analyse; if not then AF2.
4. System reads the amount of liquid on the plate; if insufficient liquid to perform the analysis then AF3.
5. Analyser performs the analysis if still allowed to analyse; if not then AF4.
6. System reads the amount of liquid on the plate; if the analysis was not performed satisfactorily then AF5.
7. Signal analysis_ok.

Failure reports:

- AF1: No plate is present in the Analyser.

  - Remedy: User loads the Analyser manually and resumes or aborts.
- AF2: Current_pos is not the upper end position to Analyse.
  - Remedy: User moves the Analyser manually to upper end position and resumes or aborts.
- AF3: There is insufficient liquid on the plate.
  - Remedy: User determines the cause of the situation and resumes or aborts the calling protocol. In case of resuming: user manually loads liquid and resumes protocol.
- AF4: Erroneous analysis command. Protocol failure.
  - Remedy: Check and change protocol or abort.
- AF5: Failure in analysis, the analysis did not use the right amount of liquid.
  - Remedy: User initiates a maintenance procedure to fix the pump precision. After maintenance user resumes or aborts protocol.

Subroutine *Deliver* captures the delivering of a plate to the Robot, provided the operating table is at its lower end position and there is a plate in the Analyser.

Typical course of events:

1. Analyser receives command *deliver* from protocol.
2. System checks if Robot is ready for plate; if not then DF1.
3. System checks if current_pos = lower end position to Deliver; if not then DF2.
4. System checks if a plate is present; if not then DF3.
5. Robot unloads the Analyser.
6. System checks if plate has been delivered; if not then DF4.
7. Signal delivery_ok.

Failure reports:

- DF1: Robot is not ready to unload the Analyser.
  - Remedy: Wait a while for the Robot to become ready.
- DF2: Current_pos is not the lower end position to Deliver.
  - Remedy: User identifies the cause and moves the Analyser manually.
- DF3: Plate is not present.
  - Remedy: User identifies the cause, and loads the Analyser manually or aborts
- DF4: The Analyser has not been unloaded.
  - Remedy: User identifies the cause, unloads the Analyser manually and resumes or aborts.

Failures:

From the alternative courses of events above we can derive the following failures:

a) Analyser move failure

b) No plate present

c) Plate already present

d) Analysis failure

e) Robot rotate failure

f) Robot arms move failure

g) Robot arms gripper failure

h) Protocol failure

Upon having this reasonably grained use case based specification, the next step is to determine the logically related use cases, and group them together. Each group of related use cases defines (informally still) a control system *component*.

3.1.2.2 Component-Oriented Development

Component-oriented development consists of reusing existing components, developing new components, and assembling new systems from them. A component is usually defined as a black box implementing a set of services, provided that the required set of context dependencies are available. A component is thus an abstraction over the space of services of a system. The services that a system has to perform, captured by the use case diagram, are partitioned into groups, each group of services being eventually implemented by a component. Components interact with each other by using the provided services. Thus, an important issue regarding a component-based system is the interactions among the participating components as shown in Figure 3.3.

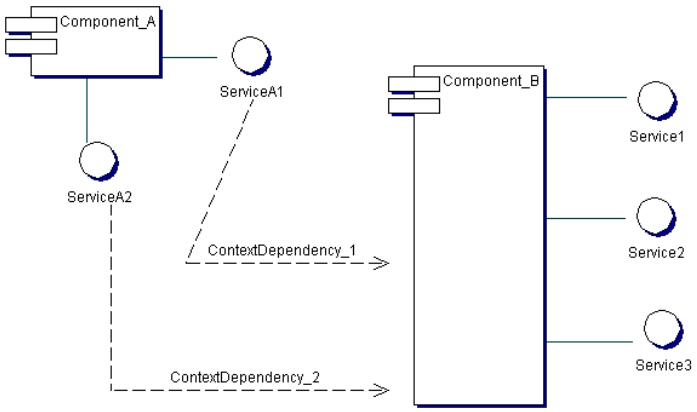


Figure 3.3: Typical component-based system.

By defining a set of components of the final system we distribute the computation already in this early development stage. The strategy of separating computation and communication from the beginning provides good support for reuse and scalability. A subset of the eventual system components may already exist and can be reused. Each of the unimplemented components is a

self-contained system of finer granularity than the initial system. Therefore, the process of specifying a system is reduced to specifying each defined, but not yet implemented component, encompassing the communication means with other components. The resulting specification models the necessary conceptual model of the component required for providing the component services.

The component diagram is deduced from the use case diagram, and the component interactions are deduced from the use case relationships. Each use case can be mapped to a component service. While the use case diagram shows only the required services and their relationships, the component diagram distributes the services to a set of components that will implement them. The conceptual model of each component, required for providing these services, is typically specified in a precise manner, using a certain specification language. We use the B-action systems to express the component specifications. Each component is modelled as one or more machines within the B Method. The services of the component are operations in the machine.

In our case study the functional requirements of the Analyser are to receive plates which are being analysed and then delivered to an external device, the Robot. From the use case diagram in Figure 3.2 we can construct an associated component diagram as given in Figure 3.4. We are only interested in the Robot loading and unloading the Analyser. The remaining services of the Robot are only given as 'other services'.

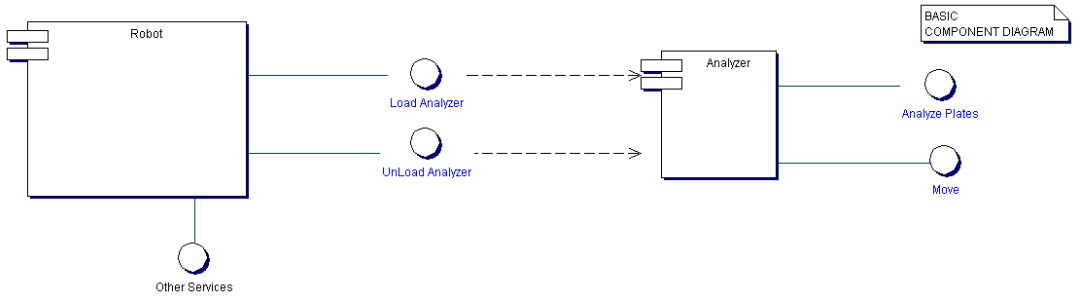


Figure 3.4: A component diagram for the analysing device.

3.1.2.3 Class Diagrams

At this level, the conceptual model of the component is shown as active classes, *i.e.*, classes whose instances have autonomous behaviour. For each class we give the attributes and their types as well as the methods. Hence, the classes model the static behaviour of the component. In the most abstract class diagram we merely consider the attributes modelling the state and the command of the component. The methods consist of the main functionality of the system and the abstract representation of the classes of errors as well as the possibility to fix these errors.

An abstract class diagram for the Analyser is given in Figure 3.5. The methods of the Analyser are the services *move*, *receive*, *analyse* and *deliver* in the use case diagram together with their possible errors and fixes. The class *ANALYSER* is the main class representing the actions of the Analyser component. The classes, *A_PROC* and *R_PROC*, represent encapsulations of the interface variables for the analyser component and the robot component. They contain procedures that can be called within the guards and actions of the analyser to set and determine the state of the interfaces between it and the robot component. The class, *GLOBALVAR_PLATE* performs a similar role to encapsulate the state concerning the presence of a plate to be analysed.

The role of these classes is elaborated later in this document when discussing the corresponding B components. *DEF0* defines the set of commands (*ACOMMAND*) in a way that can be used by several classes/components.

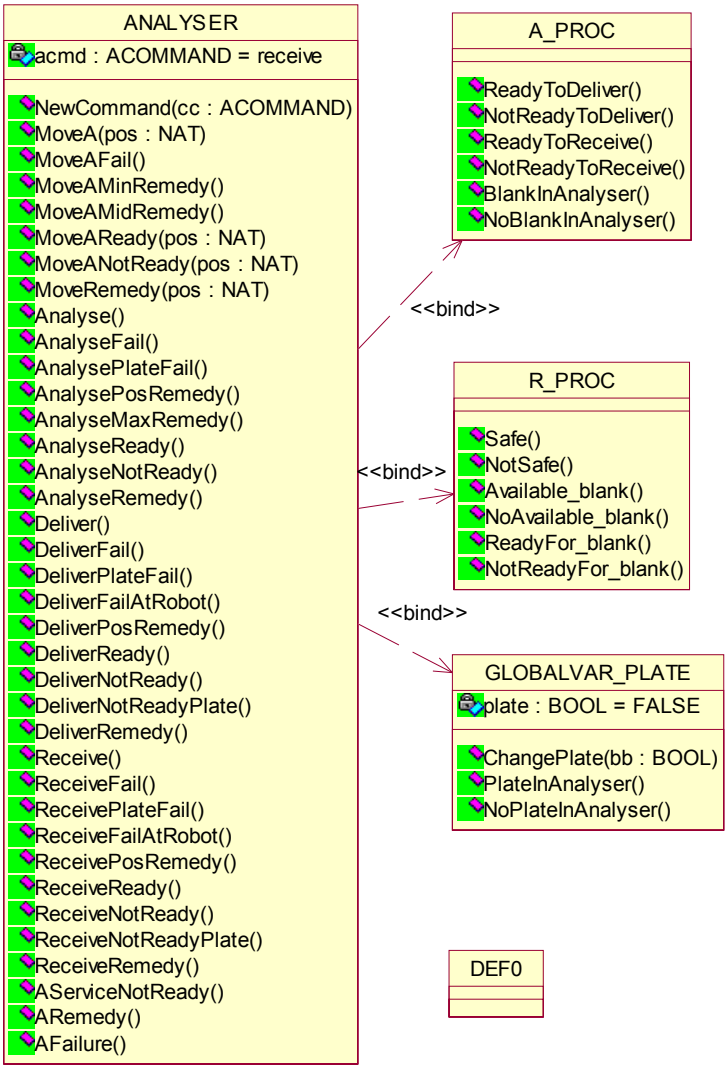


Figure 3.5: An abstract class diagram of the Analyser.

3.1.2.4 Statechart Diagrams

The autonomous behaviour of each class is specified using statechart diagrams. At least three such statechart diagrams are developed.

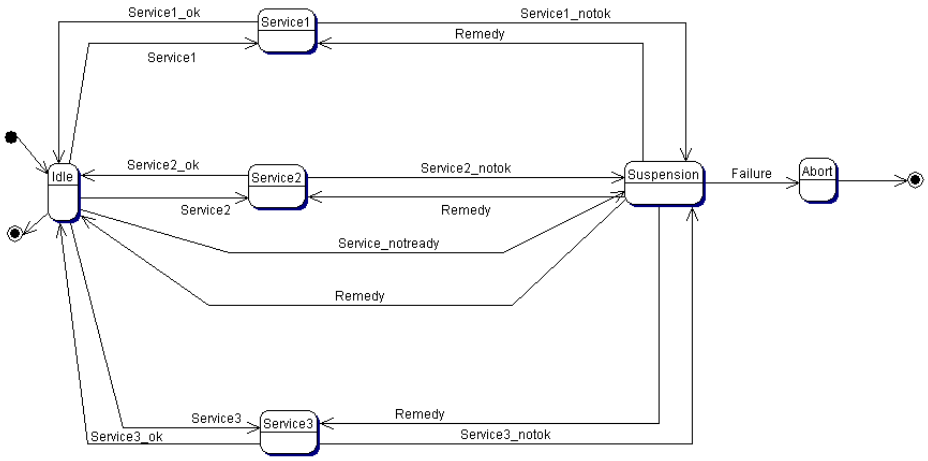


Figure 3.6: A basic statechart diagram for a component.

The first primitive statechart diagram is shown in Figure 3.6. The services in this statechart diagram are derived from the services in the use case diagrams. Here we merely model state transitions and events causing these transitions. We go from state *Idle* to a service *n* or to state *Suspension*. After service *n* is performed we return to state *Idle* or go to state *Suspension*. From state *Suspension* we may return to service *n* or to state *Idle,* or go to the state *Abort*. The value of state is analogous to an additional attribute of the class whose behaviour is being described.

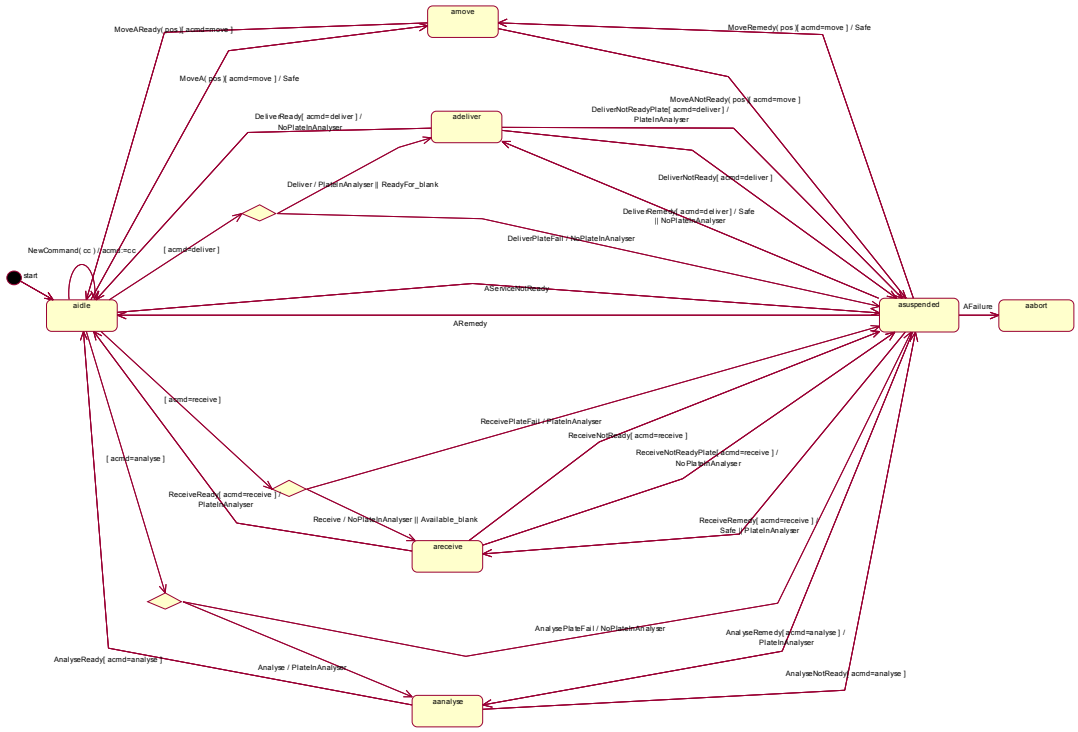


Figure 3.7: Basic statechart diagram for the Analyser.

The basic statechart diagram of the Analyser in Figure 3.7 can be derived from the use case diagram in Figure 3.2. The services *move*, *receive*, *analyse* and *deliver* in the use case diagram

form the states in this basic statechart diagram together with the states *idle*, *suspended* and *abort*. The methods of the class diagram are given as the transitions between these states. This basic statechart diagram and the class diagram in Figure 3.5 are translated to B with the U2B-tool [Snook00].

3.1.2.5 Creating a B Model from UML

In order to be able to ensure the correctness of the system we need a formal analysis tool. A formal method that comes with such tools is the B Method. In order to be able to reason about distributed systems within the B Method we use B-action systems [Walden98a]. Event B also supports the development of distributed systems and indeed the extensions of event B are built on action systems. However, some features for action systems like procedure handling are not supported in event B yet. Procedures are important in distributed systems, since they provide more structuring to the systems, as well as a general communication mechanism between interacting systems. Hence, we apply action systems within B to model the distributed systems within the healthcare case study.

```
MACHINE Component
VARIABLES
   state, cmd
INVARIANT
   state ∈ { Idle,Service1,Service2,Service3,Suspension,Abort} ∧
   cmd ∈ {serv1,serv2,serv3}
INITIALISATION
   state := Idle || cmd :∈ {serv1,serv2,serv3}
OPERATIONS
   New_command(ss) =
      PRE ss ∈ {serv1,serv2,serv3} THEN
         SELECT state = Idle  THEN cmd := ss END
      END;

   Service1 =
      SELECT cmd = serv1  ∧  state = Idle THEN state := Service1 END;

   Service1_fail =
      SELECT cmd = serv1  ∧  state = Idle THEN state := Suspension END;

   Service1_ok  =
      SELECT state = Service1  THEN state := Idle END;

   Service1_notok =
      SELECT state = Service1 THEN state := Suspension END;
   …
   Service_notready =
      SELECT state  = Idle THEN state := Suspension END;

   Remedy =
      SELECT state = Suspension
      THEN state :∈ {Idle,Service1,Service2,Service3} END;

   Failure =
      SELECT state = Suspension THEN state := Abort END
END
```

Figure 3.8: An abstract B-action system.

The first step in our formal development is to create an abstract B-action system from the basic statechart diagram in Figure 3.6. The tool U2B [Snook00] supports this translation. The event triggering state transitions (names assigned to arrows) correspond to operation names in the abstract machine specification in Figure 3.8. Each service of the basic statechart diagram will have four operations: the first for checking its precondition, the second for taking care of a false precondition, the third for executing the proper result of an operation and the fourth for taking care of a failed computation. We also have an operation `Remedy` for fixing an exception and continuing the computation, and an operation `Failure` taking the system to the state *Abort* from where no further execution is possible. The attribute `cmd` corresponding to *ServiceN* in the statechart diagram as well as the attribute `state` form the variables of the abstract machine `Component`.

As an example of the U2B translation we show the part of specification of the Analyser in Figure 3.9 (the complete machine ANALYSER is presented in Appendix A.1).

```
MACHINE ANALYSER
SEES  DEF0
INCLUDES
    GLOBALVAR_PLATE,
    R_PROC,
    A_PROC
SETS
    ASTATE = {aidle,amove,aanalyse,arecieve,adeliver,aabort,asuspended}
VARIABLES
    astate,
    acmd
INVARIANT
    astate ∈ ASTATE ∧
    acmd ∈ ACOMMAND
INITIALISATION
    astate := aidle ||
    acmd := receive
OPERATIONS
   NewCommand (cc) =
      PRE
         cc ∈ ACOMMAND
      THEN
         SELECT astate=aidle THEN acmd:=cc END
      END;
   MoveA (pos) =
      PRE
         pos ∈ NAT
      THEN
         SELECT astate=aidle ∧ acmd=move
         THEN   astate:=amove || Safe
         END
      END;
   MoveAFail  =
      BEGIN
         SELECT astate=aidle ∧ acmd=move
         THEN   astate:=asuspended || NotSafe
         END
      END;
   …
END
```

Figure 3.9: B machine of the ANALYSER specification.

The operations in a B-action system consist of guards and actions as in event based B. We can also have global variables that can be read and updated by more than one system. The global variable is declared in a separate machine `Global_z` which is *included* in the communicating systems `A_comp` and `B_comp`. By including the machine `Global_z` in `A_comp` and `B_comp`, they are both allowed to assign a value `v` to the variable `z` via the operation `assign_z(v)` [Butler96].

```
MACHINE  A_comp
INCLUDES  Global_z
VARIABLES
   …
END
```

```
MACHINE  B_comp
INCLUDES  Global_z
VARIABLES
   …
END
```

```
MACHINE  Global_z
VARIABLES
  z
INVARIANT
  J(z)
INITIALISATION
  z := z0
OPERATIONS
   assign_z(v) = PRE J(v) THEN z := v  END
END
```

Previously developed machines that have been shown correct can be used in the development by including them in the component using `SEES`, `INCLUDES` or `EXTENDS` [Abrial96]. The procedures declared in a B-action system can be local or global. The local procedures are declared and referenced locally within the same system. The global procedures, on the other hand, may be referenced by other B-action systems as well. The global procedures are services provided/claimed between systems (components) and can be derived from the component diagram. Let us consider the exported procedure `P` declared in system `B_comp` and called from system `A_comp` as shown below.

```
MACHINE  A_comp
INCLUDES  B_comp
VARIABLES
   x
INVARIANT
  I(x)
INITIALISATION
 x := x0
OPERATIONS
   A1 = SELECT Q1 THEN T1 || P END;
   …
END
```

```
MACHINE  B_comp
VARIABLES
  y
INVARIANT
  J(y)
INITIALISATION
  y := y0
OPERATIONS
   P = SELECT G THEN U END;
   …
END
```

We note that the B method requires the calling operation and the procedure to be in separate machines. Hence, in case an operation in `B_comp` also calls the procedure `P`, we have to outsource `P` into a separate machine and incorporate them via the `INCLUDES` clause. When operation `A1` calls procedure `P`, `A1` is enabled only if procedure `P` is also enabled, *i.e.*, `Q1` $\wedge$ `G` holds. The procedure `P` and the operation `A1` are executed as a single atomic entity. The procedures are discussed in more detail elsewhere [Sere00, Walden98b].

Formal Specification of the Analyser

From the class and the statechart diagrams in Figure 3.5 and Figure 3.7. We can get the abstract B-action system `ANALYSER.mch` automatically with the tool U2B. The Analyser has four services/subroutines *move, receive*, *analyse* and *deliver*. The only variables in this first machine are the variables modelling the states `astate` and the current command `acmd`, as well as a global variable `plate` stating whether there is a plate in the Analyser or not. The global variable is given in the machine `GLOBALVAR_PLATE.mch` in Appendix A.3.

The global procedures of the Analyser are given in the machine `A_PROC.mch` in Appendix A.4. The global procedures of the Robot are the imported by the Analyser. The Analyser has the global procedures `ReadyToDeliver`, stating if the Analyser is ready to deliver a plate or not, `ReadyToReceive`, stating if the Analyser is ready to receive a plate or not, as well as `BlankInAnalyser` stating whether the Analyser senses a plate or not. In order for the Analyser to be able to function correctly it checks via the imported procedures whether the Robot is safe or not, `Safe`, whether the Robot has a blank available or not, `Available_blank`, and whether the Robot is ready to receive a blank or not, `ReadyFor_blank`. The global procedures are only modelled as *skip* statements here, but they will be refined later in the development.

The Robot has to be in a safe state when the Analyser is moving and the plate has to be in the Analyser during an analysis. During the delivery operation the Robot should also be ready to receive the plate that should exist in the Analyser, while the receive operation requires the Robot to have a plate available for the empty Analyser. A new command can be given each time the Analyser is in state `idle`.

3.2 Computation Models

3.2.1 Modelling a Byte Code Verifier

In this section, we describe the methodology used to develop models of the type verifier and structural verifier. It is important to note that a formal specification cannot be constructed out of nothing; the goals of the verifier need to be established, before it can be modelled. Hence, the development of an informal specification is an important mandatory step that simplifies the model construction.

This section focus on the methodology used to translate informal specifications into formal ones. We distinguish between two cases: the type verifier and the structural verifier, because they require two different modelling techniques.

3.2.1.1 Modelling a Byte Code Instruction

Despite the modelling of the two nested loops, the one over the methods of the package being verified and the one over the instructions contained within a particular method, the heart of a type verifier is the modelling of each byte code instruction. Note that the Java Card language contains 184 different byte codes. In this section, we present the methodology that we have chosen in order to construct the formal specification of each byte code instructions. We can identify three steps to model a byte code instruction as shown in Figure 3.10. First, pertinent information is extracted from the requirements laid down by Sun. Then, information is gathered into an informal document describing the typing behaviour of each byte code instruction. Finally, this informal description into a formal specification.

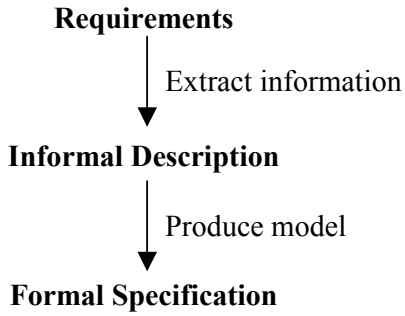


Figure 3.10: Building a computational model.

Extracting Information from the Requirements

The first step in modelling an instruction is to understand the Java Card language. [Sun00] contains a description of the behaviour of each instruction of the Java Card Virtual Machine (JCVM). The main problem with this informal specification is that typing and operational information is mixed. The goal of the first step is to determine the principles of the Java Card and moreover the language itself. This comprehension helps to distinguish between static and dynamic semantics.

aaload

Load reference from array

Stack

…, *arrayref*, *index* .

…, *value*

Description

The *arrayref* must be of type reference and must refer to an array whose components are of type reference.The *index* must be of type short. oth *arrayref* and *index* are popped from the operand stack. The reference *value* in the component of the array at *Index* is retrieved and pushed onto the top of the operand stack.

Runtime Exceptions

If *arrayref* is null, *aaload* throws a NullPointerException. Otherwise, if *index* is not within the bounds of the array referenced by *arrayre*f, the *aaload* instruction throws an ArrayIndexOutOfBoundsException

Figure 3.11: Sun's specification of the `aaload` instruction.

The example provided in Figure 3.11 gives the semantics of a JCVM instruction called `aaload`. In the *Description* paragraph, the two first few sentences describe type verification and the two last sentences describe the dynamic JCVM execution. The *Stack* paragraph describes the stack evolution. The *Exceptions* paragraph is not really useful, one just has to know that this byte code can throw exceptions. It is important to note that in the *stack* paragraph, Sun provides values whereas we want, for the type verifier, the type of elements in the stack. So there needs to be some interpretation of the specification.

Gathering Information into an Informal Description Document

From the specification provided by Sun, given in Figure 3.11, the person who writes the informal specification of the type verifier needs to extract information and to format it in a suitable way. To do so, we have proposed a format containing 5 different parts:

- The first part describes the stack from a typing point of view. In fact, it corresponds to what is expected by the instruction to execute itself and how the stack is transformed. In Figure 3.12, it says that there are at least two elements in the stack. If the stack contains a pointer to an array of objects and a short, then the instruction returns the contents of the array at the index specified by the short. Otherwise, if the stack contains a `null` pointer and a `short`, the instruction returns a `null` pointer.
- The second part of the specification concerns the pre-modification tests. In fact, it consists in giving the tests that have to be verified before executing the instruction. In the example in Figure 3.12, these tests check that there are at least two elements in the stack and that these two elements have a correct type, according to the static semantics.
- The third part describes the modifications performed by the instruction. The modifications concern the evolution of the stack or of the local variables. In the example provided in Figure 3.12, the two elements at the top of the stack are removed and an element is pushed onto the stack. The type of this latter element is function of the type of the second element in the stack before the modification. Note that in this case, local variables are not modified but others instructions modify them.
- The fourth part contains tests that allow one to ensure properties of the modification performed by the instruction. In the example, there is are post-modification tests. Finally, the last part indicates the categories of exceptions that can be thrown by the instruction. It is necessary to know if the instruction can throw an exception or not for the modelling.

aaload

[…, `refarray class`, `short`] => […, `ref class`]

[…, `null`, `short`] => […, `null`]

Pre-modification tests:

1. The stack must contain at least two elements **(*1*)**
2. The two topmost elements of the stack have to be of types compatibles with refarray `class` and `short`. **(*2*)**

Modifications: **(*3*)**

The two topmost elements of the stack are removed.
If the second element was a `refarray` type, then a reference of the same class is pushed onto the stack. Otherwise a type `null` is pushed.

Post-modification tests:

None

Throws

- `NullPointerException`
- `ArrayOutOfBoundException`
- `SecurityException`

Figure 3.12: Informal specification of the `aaload` instruction for the verification.

Each instruction of the byte code is described following this template in an internal document. This document represents the requirements that help to define the type verifier. This internal document is then the only reference used by the formal developer to translate an informal description into a formal specification of the type verifier. Note that the writer of this informal description is not the same person as the formal developer. The main reason is that in this way,

the formal developer, when modelling the type verifier, asks the writer of the informal description for more information. Thus, the formal developer performs a review of the informal description. It enables ambiguities to be removed from within the informal document and the production of a clearer informal specification and a more precise formal specification.

Translating Informal Description into a Formal Specification

The last step in the modelling of the type verifier is the translation from the informal description to a formal B specification. This stage is critical , as the final code and the proof are constructed from this formal specification. Therefore, if there is a mistake in the translation, the proof will not be able to detect it. Moreover, the proof may be useless since the model will not reflect the correct behaviour of a type verifier.

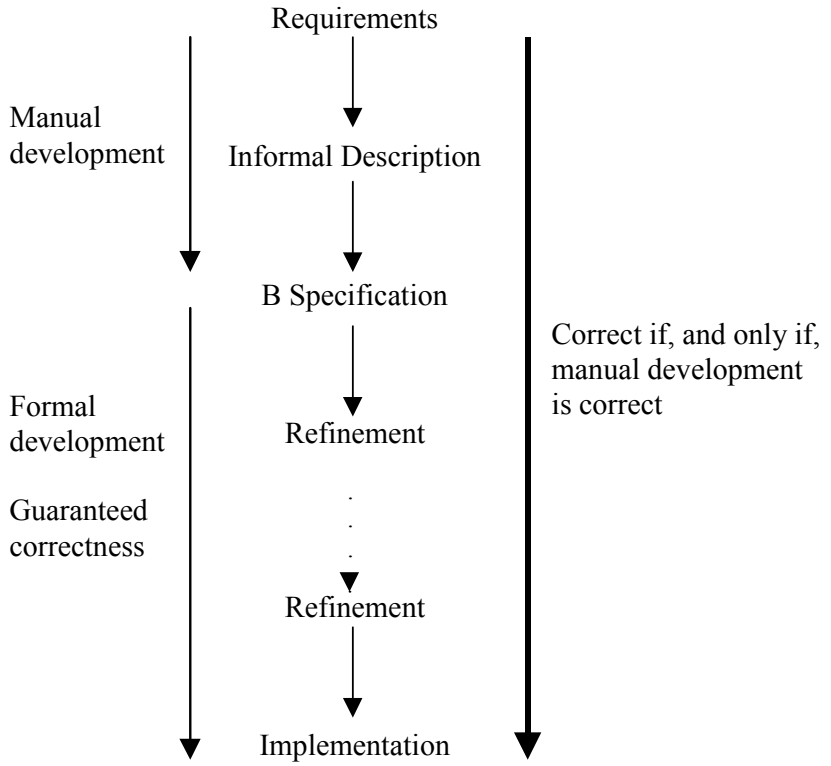


Figure 3.13: Steps for building a formal model.

```
bb ← verify_aaload =
BEGIN
   SELECT
      2 ≤ size(stack) ∧                                          (1)
      last(stack) = c_short ∧                                    (2)
      last(front(stack)) = c_refarray ∧
      (pc ∈ dom(exception_handler)                               (4)
      ⇒
      ∀label.(label ∈ exception_handler(pc)
```

```
        ⇒ COMPATIBLE(loc_var, loc_var_descriptor(label))) ∧
        c_uref ∉ ran(loc_var))
    THEN
        bb := TRUE ||
        stack := front(front(stack)) ← c_ref                    (3)
    WHEN
        2 ≤ size(stack) ∧                                       (1)
        last(stack) = c_short ∧                                 (2)
        last(front(stack)) = c_null ∧ (pc ∈ dom(exception_handler)   (4)
        ⇒
        ∀label.(label ∈ exception_handler(pc)
            ⇒ COMPATIBLE(loc_var, loc_var_descriptor(label))) ∧
            c_uref ∉ ran(loc_var))
    THEN
        bb := TRUE ||
        stack := front(front(stack)) ← c_null                   (3)
    ELSE
        bb := FALSE
    END
END
```

Figure 3.14: Formal specification for `aaload` instruction verification.

Figure 3.14 concludes the example of the `aaload` instruction that we have chosen. In fact, we use this example to study the correspondence between the informal description and its B abstraction. We use italic bold numbers on the figures to show the correspondence between sentences of Figure 3.12 and the B specification in Figure 3.14. Hence, ***(1)*** indicates that there must be at least two elements on the top of the stack. ***(2)*** shows that the element on top of the stack must be a `short` and that the second element is either a reference to an array (*c_refarray*) or a `null` pointer.

As shown in Figure 3.12, we have to distinguish these two cases because the `aaload` byte code instruction has two potential treatments, the `SELECT` clause is the most suitable. This clause allows us to specify an action guarded by a condition. In this example and all the other instructions, all the `SELECT` clauses are deterministic and the global specification is deterministic. The operation returns a result, a value is given to the variable containing the result in each branch and the stack is modified in the two correct branches (two elements are popped and one is pushed).

In the `SELECT` and `WHEN` clauses, the developer indicates the pre-conditions that allow the execution of the instruction. In our example, the pre-conditions state that there must be at least two elements in the stack and that the top element is a `short` and the second element is either a reference to an array of objects or a `null` pointer. We also add the test for the exception handler ***(4)***. In fact it is an additional condition which checks that if the current program counter *pc* is in an exception handler, we have to check that for the proof label given for each handler: the local variable is compatible with the local variable descriptor associated with the label, and that there are no uninitialised references in the local variables. This test is factored out in the internal document describing the requirements for the type verifier.

Finally, in the `THEN` and `ELSE` clauses, we indicate the behaviour of the instruction. In both cases, when dealing with non-exception behaviour, the return variable is set to `TRUE` and the stack is modified consequently to the test. Two elements are popped from the stack and a reference is pushed onto the stack. In the case of exceptional behaviour, which is caused by there being insufficient elements on the stack, the typing of the elements is incorrect or the exception

handler checks are incorrect. Here, the instruction returns `FALSE` and does not modify the stack or the local variables.

If an exception is raised, it is propagated within the program and the verification process is stopped. Otherwise, the verification process proceeds with the next instruction.

3.2.1.2 Modelling a Structural Test

If the type verifier appears to be a single, complex machine that ensures the enforcement of the typing rules, the structural verifier appears more as a collection of tests that have to be checked in order to accept data. In this section, we describe how we construct these tests by providing the methodology used, the templates and some examples. The structural verifier ensures that data received by the card is in the correct format, *i.e.*, the format defined by Sun in [Sun00]. To do so, we have to model each component of a CAP file by representing their properties and checking that data which is supposed to correspond to a component does indeed do so.

As for the type verifier, we have developed a methodology to model each component. We describe this methodology in this section and discuss the obtained result.

Extracting Information from Requirements

In [Sun00], Sun informally specifies the CAP file format. As the goal of a structural verifier is to ensure that a given set of data is well formed, this informal specification is the basis. However, information strictly related to structural verification has to be extracted from this informal specification. In fact, Sun's documentation not only includes an explanation of what elements can be found in a specific component, but also information not related to it. Figure 3.15 provides an example of Sun's specification. It concerns the definition of the `Export` component, one of the eleven components of the CAP file defined by Sun. In this description, Sun provides a description in a C-like notation for the structure of the component and then a description in a natural language for each of the elements. In particular, the C-like notation includes the size of each element given with `u1` or `u2` indicating whether the element is a byte or word, respectively. In this definition of a component, the size of each element is needed for the structural verification. The export component is represented by the following structure:

```
export_component {
       u1 tag
       u2 size
       u1 class_count
       class_export_info {
          u2 class_offset
          u1 static_field_count
          u1 static_method_count
          u2 static_field_offsets[static_field_count]
          u2 static_method_offsets[static_method_count]
} class_exports[class_count]
}
```

tag

The `tag` item has the value COMPONENT_Export (10).

size

The size item indicates the number of bytes in the export_component structure, excluding the tag and size items. The value of the size item must be greater than zero.

class_count

The class_count item represents the number of entries in the class_exports table. The value of the class_count item must be greater than zero.

class_exports[]

The class_exports item represents a variable-length table of class_export_info structures. If this package is a library package, the table contains an entry for each of the public classes and public interfaces defined in this package. If this package is an applet package, the table contains an entry for each of the public shareable interfaces defined in this package. An index into the table to a particular class or interface is equal to the token value of that class or interface (§4.3.7.2). The token value is published in the Export file (§5.7) of this package.

The items in the class_export_info structure are:

class_offset

The class_offset item represents a byte offset into the info item of the Class Component (§6.8). If this package defines a library package, the item at that offset must be either an interface_info or a class_info structure. The interface_info or class_info structure at that offset must represent the exported class or interface. If this package defines an applet package, the item at the class_offset in the info item of the Class Component must be an interface_info structure. The interface_info structure at that offset must represent the exported, shareable interface. In particular, the ACC_SHAREABLE flag of the interface_info structure must be equal to 1.

static_field_count

The static_field_count item represents the number of elements in the static_field_offsets array. This value indicates the number of public and protected static fields defined in this class, excluding final static fields of primitive types. If the class_offset item represents an offset to an interface_info structure, the value of the static_field_count item must be zero.

static_method_count

The static_method_count item represents the number of elements in the static_method_offsets array. This value indicates the number of public and protected static methods and constructors defined in this class. If the class_offset item represents an offset to an interface_info structure, the value of the static_method_count item must be zero.

static_field_offsets[]

The static_field_offsets item represents an array of 2-byte off-sets into the static field image defined by the Static Field Component (§6.10). Each offset must be to the beginning of the representation of the exported static field. An index into the static_field_offsets array must be equal to the token value of the field represented by that entry. The token value is published in the Export file (§5.9) of this package.

static_method_offsets[]

The static_method_offsets item represents a table of 2-byte offsets into the info item of the Method Component (§6.9). Each offset must be to the beginning of a method_info structure. The method_info structure must represent the exported static method or constructor. An index into the static_method_offsets array must be equal to the token value of the method represented by that entry.

Figure 3.15: Sun's specification of the Export component.

In this description, there is also information about other components. For example, in Figure 3.15, the class_offset element represents a memory offset into the Class component. Therefore, interactions between components are possible. The remaining information that appears in this description concerns information not related with verification.

This description contains most of the information necessary to perform internal and external structural tests. However, as we would like to differentiate between internal and external tests as in the architecture, we have decided to distinguish them. Therefore, we have to write a specific internal document, as for the type verifier, containing the different tests by component.

Gathering Information into an Informal Description Document

From the description provided in Figure 3.15, the person writing the informal requirements of the structural verifier can extract tests related to structural verification. This document is written in natural language but tries to classify the tests into two categories per component: the internal

tests containing the format of the component, and the external tests checking information shared by several components.

Figure 3.16 and Figure 3.17 provide examples of tests extracted from Sun's specification. For example, to check that the `Export` component is well formed, one has to verify that the `tag` is equal to `10`, that the table `static_field_offsets` has `static_field_count` entries and the table `static_method_offsets` has `static_method_count` entries.

- Check that `Tag = 10`
- Check that the number of entries in the `static_field_offsets` is equal to `static_field_count`.
- Check that the number of entries in the `static_method_offsets` is equal to `static_method_count`.

Figure 3.16. Informal specification of internal tests related to Export component.

- Check that the value of `component_size[tag-1]` of the Directory component is equal to the `size` item: `size = component_size[tag-1]`.
- Check that the `ACC_Export flag` of the `flags` item in the `Header` component is equal to `1`.

Figure 3.17: Informal specification of external tests related to Export component.

For all the components specified by Sun, and an additional one for the *proof* component for the PCC (Proof Carrying Code) technique, there is a correspondence in the internal document that we have written. In fact, the internal document specifying the informal requirements for the structural verifier is organised as depicted in Figure 3.16 and Figure 3.17. The document produced can now be used by the developer of the structural verifier.

Modelling a Component

The first step in the structural verifier is to provide a B specification of each CAP file component. This first B abstraction, describes what information needs to be by each component. As the CAP file is composed of eleven standard components, we aim to model each component in a similar way in order to split the development between different developers. Therefore, we propose the template in Figure 3.18. This template represents the abstract B machine. The goal of this template is to simplify the B development. It also includes a naming convention for the machines, the variables and the operations of the machine. It is important as the proof of each component can be made independently; changing the name of a variable or an operation generally implies that the proof must be restarted, almost from the beginning.

In the `VARIABLES` clause of the specification, one adds the variables needed to describe the components. In the case of the `Export` component, this clause includes *exp_class_count, exp_class_offsets* etc. that correspond to elements of the corresponding informal specification (Figure 3.15). Moreover, we define a specific variable named *component_verified*, *i.e.*, *exp_verified* in the case of the *Export* component, that indicates whether this component has been successfully verified or not. At the beginning, this variable is set to FALSE.

In the `INVARIANT` clause of the specification, one indicates the properties of the variables defined in the `VARIABLES` clause. It contains the typing of each variable and the properties they must respect. For example, in the case of the `Export` component, one property is that the *exp_class_count* is a `byte` (*exp_class_count* $\in$ *t_byte*) and that its value is strictly positive (*exp_class_count* > 0). In the `INVARIANT`, the variable indicating that the verification of the

component has been performed is typed as a boolean. This clause also includes the property that is enforced when the component is verified.

In the INITIALISATION clause, variables are initialised for the first time. Note that *component_verified* is set to FALSE, indicating that, at the beginning, the component is not verified.

```
MACHINE cpn_component

VARIABLES
   <Set of variables used to describe the component>,
   Component_verified

INVARIANT
   <Set of properties on variables previously defined> ∧
   component_verified: BOOL

INITIALISATION
   <Initialisation of all variables describing the component> ||
   component_verified := FALSE

OPERATION
   Res ← component_internal_verif =
      PRE component_verified = FALSE
      THEN   component_verified = TRUE ⇒ <the component is correct>
      END;
   Res ← other_services_1 =
      PRE component_verified = TRUE
      THEN   …
      END
   …
END
```

Figure 3.18: Template of the model of a CAP file component.

Finally, the last clause is that of OPERATIONS. This clause contains all the services of the machine. We distinguish two kinds of services: first an operation is declared called *component_internal_verif*. This operation performs the internal verification of the component. In order to be called only once, this operation is pre-conditioned by the *component_verified*. That is, if this variable is FALSE, then the operation can be executed otherwise, it has already been executed (and the verification is a success) and there is no need to execute it again. This particular operation contains the properties which ensure the correctness of the component. The refinement of this operation is performed by implementing the list of internal tests proposed in the internal document written using the Sun specification and of which some examples are provided in Figure 3.16. The second kind of service concerns operations allowing access to the content of the component. These operations are pre-conditioned with the same variables but they can be executed only if the component has been verified, *i.e.*, *component_verified* is TRUE. These pre-conditions ensure that data is not accessed unless the component has been verified and that it then can be considered as a valid component, from a syntactic point of view.

This template provides a general framework for writing B specifications of each component, and makes communication between developers easier. The abstract machines obtained are then refined in order to map this abstraction with the real data made from zeros and ones. Moreover, this template helps formal developers to relatively easily translate the informal description as this formal specification is not low-level and represents the component. In fact, this B specification provides a view of each component that is used to model external tests and the type verifier.

Modelling an External Test

Modelling an external test is easier than modelling a CAP file component. The main reason is that we abstract each component to build higher level tests. In fact, external tests involve information, and so variables, that reference or represent the same value. Therefore, these tests rely on the CAP file model, on the variables defined in each of these model and their properties. So it is relatively straightforward to translate the informal specifications into the formal ones.

For example, the formal specification of the two tests provided in Figure 3.17 is depicted in Figure 3.19. The first test specifies that the size of the `Export` component is the same as the one stored in the *dir_component_sizes* table of the `Directory` component. Therefore, in B, one has just to write the test in a specific operation. Moreover, we have chosen to write one test per operation. Of course some tests can be factorised into a single operation, but allowing one test per operation makes the B machine clearer to read, write and easier to maintain. Each operation is pre-conditioned by tests. These tests, in our case *exp_verified* and *dir_verified* for the first precondition, ensure that both the `Export` and the `Directory` component have been verified before any external test can be made on them. The body of the operations, contain the properties that need to be checked. The operation returns the result of the test: `TRUE` if the test passes, `FALSE` otherwise.

```
res ← exp_ext_1 =
    PRE
        exp_verified = TRUE ∧
        dir_verified = TRUE
    THEN
        res := bool(exp_size = dir_component_sizes(c_tag_export - 1))
    END;

res  ← exp_ext_2 =
    PRE
        exp_verified = TRUE ∧
           hdr_verified = TRUE
    THEN
        res :=  bool(hdr_has_export = TRUE)
    END;
```

Figure 3.19: B specification of external tests for the `Export` component.

The second test modelled involves the `Export` and the `Header` components. The property checks that, an element indicating the presence of the `Export` component in the `Header` component is set to `TRUE`. Note that it is still a simple property to express as the developer of the external tests does not have to model the whole problem. In fact, most of work has been done whilst modelling the component. The developer now has access to a formal description of each component and just has to express the relationships between the variables.

Embedding External Tests into a B Component

As we have just seen, the specification of external tests is simple. It is a translation of the informal specification into a formal specification. In order to simplify the B development and to have the same format for the B specification for each external test, Figure 3.20 gives a template for writing external tests. First, we consider that each component *cpn_component* has a machine *cpn_component_ext* that contains the external tests of this component on the other components.

```
MACHINE
    cpn_component_ext

SEES
    <All cpn_components concerned by the consistency of the component>

OPERATIONS
    Res ← test1 =
        PRE component_verified = TRUE ∧
            component1_verified = TRUE ∧
            …
        THEN
            Res := bool( <Description of the property> )
        END

    Res ← test2 =
        PRE …
        THEN …
        END;
    …
END
```

Figure 3.20: Template for modelling the external tests for the structural verifier.

The SEES clause of the B machine lists all the components involved in tests described within this machine. It contains the names of components interacting with the component being tested, and some context machines that provide definitions. There is no need to add any invariants since the proof relies on checking the consistency of component invariants and their properties.

There is no need to have an INITIALISATION clause either, since all variables are already initialised, as they belong to other machines.

So, the final B clause is that of the OPERATIONS. Each operation describes one or more tests. The pre-conditions statement ensures that internal verification of each component affected by the test has been performed. This allows the B prover to obtain more properties and hypotheses. The body of the operation contains the formal specification of the test, *i.e.*, the abstract representation of what is performed by this test. In this representation, abstract variables are used to express the property of the specified test. That is, the description of the relationship between abstract variables of the different components affected by the test (as shown in Figure 3.19). The next step in the refinement process is the implementation. In the implementation, all abstract variables are replaced by concrete ones and access to data is performed by the different operations of each component.

4 Analysing Models

This chapter describes different ways in which formal methods may be analysed for validity and soundness.

4.1 Modelling in B

4.1.1 Quick Rules

The following commandments have been successively used on significant projects:

1. Before constructing the AMN model, discover and arrange its most important entities within a not too constraining framework.

2. Adopt a strict top-down approach to analyze the problem and construct the AMN model.

3. Use abstract machines, but also fairly simple implementations, to define an abstract, hierarchical, and modular AMN model of the system.

4. Introduce critical properties in invariants or operations as high as possible in the hierarchical AMN model.

5. Construct small invariants and operations. To do it, introduce intermediate abstract machines that break down large invariants and operations.

6. Before proving and refining a software model, check the relevance and the tractability of the abstract machines, of their interfaces, and of the architecture. To do it, construct the complete architecture of the AMN model before proving and refining.

7. Avoid under-specification, even if it seems redundant, first define completely the desired feature in abstract terms, and then implement it concretely.

4.1.2 Modelling

This chapter aims at providing some useful hints on the modelling with B. These hints are a mix of organizational and technical recommendations.

4.1.2.1 Determining Properties

Method

With a clear and complete statement of requirements, it is possible to start the construction of a system. This is the role of the system specification phase. The designer analyses the problem, and elaborates an abstract model of it, not an abstract solution to it. Later on, s/he identifies the functions and the data that fulfil the requirements, determines their arrangement, and gives a structured description adequate for the forthcoming formalization.

It is a top-down process which goes, by successive refinement and decompositions, from the most general and abstract description of the system to the most detailed and concrete. The

process often implies iterations, since information discovered at some stages of analysis may call into question the stages carried out before.

Some of the basic principles of the method are:

- To delimit precisely the system to be analysed, and to state all its links with the environment.

- To deal simultaneously with a reduced amount of information in order to overcome the complexity of the analysed system (divide and conquer approach).

- To decide of a particular point of view and to stick to it all along the analysis. Actually, systems can be analysed from different points of view (usage, safety, design, maintenance, *etc.*) and it is important to give a homogeneous description.

- To describe the analysed system in terms of activities and data arranged in a hierarchical structure that exhibits the flow of information and the dependence, causality, and decomposition relationships.

- To give a more and more detailed description as and when the analysis progresses until all the required features are captured.

- To adopt the adequate level of abstraction for the type of description built. Depending on this level, the description may tell either what the analysed system does, which is adequate for software specification, or how it does it, which is adequate for design.

According to the level of decomposition the introduced data must be global and abstract or detailed and concrete.

In the top-level component, two kinds of data are good candidates to be introduced:

- The high level data strictly necessary to determine the most important properties of the system.

- The high level data strictly necessary to formulate a major global requirement.

Properties

Statements concerning functions may be predicates defining an activation condition, predicates formulating an input/output relationship, or automata describing a behaviour. The physics of the monitored phenomena is a good source of predicates. Statements concerning data may be predicates describing invariant properties or dependence properties. The range of values of data and the equations that govern the safety conditions of the controlled system are examples of invariant properties. The relations between data of a given level of decomposition and data of a lower level which is needed to calculate the former are examples of dependence properties.

It is not necessary to formalise all the properties of the data and the functions of the system right from the beginning. This would introduce too concrete data structures, and lead to models hard to read and to prove.

Invariants

The place of invariants influences the architecture of the AMN model, for two reasons. First, since invariants usually relate variables, and since operations must preserve invariants, an operation updating a variable of an invariant must update the other variables in order to maintain the balance. Consequently, the same checks and the same updates may appear in all these operations, and they will have to be refined, implemented, and proved several times. Therefore, in order to reduce redundancy, the abstract machines that set invariants relating many variables should have few and global operations. Second, in a top-down constructed AMN model, it is the abstract machines at the upper levels of the hierarchy that have few and global operations. Therefore, it is recommended to place invariants relating several variables as high as possible in the hierarchy, in abstract machines containing few and global operations.

It may happen that systems are concerned with global safety requirements that relate two consecutive states of the system. Usually, it is painful to formalise these requirements with invariants, as it is necessary to introduce extra variables that increase the complexity of the model. On the other hand, when an abstract machine declares a single operation, it is equivalent to formalise requirements with the invariant or with the operation, since in the former case the operation preserves the invariant, while in the latter it establishes that invariant after each invocation. Moreover, the definition of an operation allows to reference previous values of the variables. So, global safety requirements that relate two consecutive states of the system, or more generally, dynamic invariants of the system, can be formalised in the definition of the single operation of the topmost abstract machine.

Invariants of an enrichment refinement must state the properties relating the introduced variables to variables inherited from the refined module.

Besides typing the variables introduced, invariants of data refinements set the relation that permits to retrieve the information stored in the replaced variables from the information stored in the replacing variables. There are distinct sorts of retrieve relations.

Sometimes, a value of the replaced variables may be represented by distinct values of the replacing variables. This is the case for example when a set of a refined module is replaced (refined) by a function: The retrieve relation indicates that the replaced set is exactly the range of the introduced function therefore, the indices of the elements of the set in the array are irrelevant, and two distinct permutations of the elements. Some other times, a value of the replaced variables is represented by one and only one value of the replacing variables.

More rarely, some values of the replaced variables are not represented by values of the replacing variables. This is possible only when the operations of the abstract module do not use effectively the non represented abstract values.

Invariants of some implementations produced at preliminary design must be completed, particularly those of the implementations that are the first and unique refinement of an abstract machine. All what has been said earlier concerning enrichment and data refinement invariants is still applicable, the concrete variables being the locally declared variables and the variables of the imported abstract machines.

Furthermore, the invariants of implementations may be used to relate the results of the operations of one imported abstract machine to the pre-conditions of the operations of another imported abstract machine. Usually, these invariants are formulated as implication predicates.

4.1.2.2 Building

Design

Design aims at identifying the different components/sub-systems and define their interfaces and their organization, that is, the implementation architecture of the system.

Completeness, simplicity, and modularity are three major concerns of design. The former deals with safety requirements capture. It determines the relevance of the model and, consequently, that of the proof, as the proof of a model which does not formalize the safety requirements of the system does not guarantee the absence of critical errors in the final system. The second concern determines the feasibility of the proof, as we have noticed that proof complexity is closely related to model complexity. Finally, the latter concern determines the organisation and the balanced distribution of the work between the development team members.

All design techniques used effectively in industry encourage the hierarchical and modular description of systems, since it is difficult, or even impossible, to give a flat description of a complex system which is clear and correct. It is necessary to introduce progressively the information and organise the processing in order to give a clear and accurate description of the system: it is difficult to construct the AMN model of an industrial application simply in terms of abstract machines. Thus, to introduce progressively the information and organise the processing the refinement technique and the programming substitutions.

The architecture of an AMN model determines the organisation of the development and the feasibility of the proof. This is why it is so important to construct a modular model with loosely coupled components which may be developed and proved separately.

Operation bodies

When operations are not defined by simple or multiple substitutions, they must be defined with “before/after” substitutions. A before/after substitution is either:

- a declaration substitution;

- or an unbounded choice substitution

Although it may be painful to define operations with “before/after” predicates, we think it is worth doing it for two reasons. First, it avoids thinking in programming terms, which is one of the most common mistakes made by B novices. Second, it forces designers to formulate operations in two different styles, and our experience tends to show that this helps to reveal either definition or implementation mistakes.

Bodies should define the final observable effects of operations on variables, they should not describe the actions carried out to obtain them. For global operations, which carry out many actions, non-determinism is of great help to fulfil this requirement.

Non-determinism is also useful to define operations that input external data in the application, as it permits to state abstractly the properties on which the application relies.

Operations of implementations should be small, they should not exceed forty lines, their average size being ten to fifteen lines.

Operations of implementations should be simple, they should avoid too many nested conditional substitutions.

The operations of enrichment refinements and data refinements are defined in the before/after substitution style, as the operations of abstract machines. The operations of algorithmic refinements are progressively constructed

Finalising preliminary design implementations means writing the implementation invariants and the loop invariants that were deliberately ignored at that phase.

Abstract Machines

It may happen that the model obtained contains too many layers involving too simple abstract machines and implementations. This makes heavy the construction and the maintenance of the AMN model, and in the worst case it penalises the performance of the eventual programs, since it generates too many nested operation calls. To avoid this situation, it is recommended to eliminate the useless intermediate abstract machines and implementations, and introduce in place data and operations of the next lower layer. In other words, it is recommended to skip over the useless abstract machines, and use in place the abstract machines that implement them.

The opposite situation may also arise, when a layer comprises too big or too complex abstract machines and implementations. Then, it is necessary either to introduce intermediate abstract machines to encapsulate and abstract parts of the data and operations of the complex components or to create local operations (operations specified and implemented in the same implementation component, that cannot be called from the outside). Local operations are also useful when building loops, as operations called within the loop do not need anymore to be replicated in imported components.

Abstract machines at the leaves of the import tree of the AMN model encapsulate the data and services on which the rest of the application is implemented. Some of these abstract machines are neither refined nor implemented formally, and their code has to be developed conventionally. They are called basic abstract machines[7].

It is good practice to group in distinct abstract machines interfaces concerning distinct devices. However, if there are not too many interfaces, it may be practical to group them all in one single basic abstract machine.

Basic abstract machines encapsulate also low level data manipulations, like data compression and decompression, interfaces of reused software components, *etc*. As any other abstract

[7] Typically, the interfaces of the runtime operating system are formally defined in one or several basic abstract machines.

machine, basic abstract machines may declare variables and should define completely their operations. However, they should not access variables of the other abstract machines of the application. That means in particular, that the exchange of data between the application and the runtime operating system must be carried out exclusively through parameter passing. Input operations return input values in their output parameters, and output operations accept output values in their input parameters. A basic abstract machine should be imported in the lowest implementation (in the hierarchy) that dominates all the abstract machines whose implementation needs that basic abstract machine, and should be seen by all the other abstract machines imported by that implementation. Thus, following our previous recommendations concerning the sees links, all the implementations that need that basic abstract machine can see it.

4.1.2.3 Verifying

Evaluation

The evaluation of the eventual AMN model obtained at preliminary design should be carried out before interactive proof[8]. As one of its aims is to check that model is well adapted for proof. The AMN model must fulfil several properties:

- It must be complete, as far as possible. Proof reading and peer review (models, variables and events dictionary) would ensure it.

- It must be relatively simple. This can be checked by looking at the number of proof obligations discharged automatically.

- Operations of implementations should be neither too big nor too complex. This can be checked also by looking at the number of proof obligations generated by the operations of the implementations.

- All the static and dynamic safety properties have been completely defined abstractly before being implemented concretely.

- Enough intermediate enrichment, data, and algorithmic refinements have been introduced. It is not recommended

- to have simple abstract machines and either a single refinement or implementation that concentrates all relevant information and produces too many or too complex proof obligations. Whenever possible, this information should be distributed among several intermediate refinements, each concerned with a particular topic.

- On the opposite, it must be checked that no superfluous refinements have been introduced that increase either the complexity or the number of proof obligations. This can be checked by suppressing the suspected refinements and comparing the complexity and the number of generated proof obligations before and after the suppression.

---

[8] Automatic proof is required as this phase provides some metrics concerning the quality of the model (in term of provability).

- Invariants of enrichment and data refinements must include all the necessary properties linking concrete variables to abstract variables. There should not be enrichment and data refinements that generate abnormally few proof obligations considering the variables of the refined module and the variables of the refinement.

Validation

In addition to conventional activities, validation of a B model consists in:

- Assessing the functional and non-functional requirements that have been taken into account in the AMN model.

- Checking that all the AMN modules have been successfully type checked.

- Checking that all the AMN modules generate proof obligations.

- Checking that all proof obligations generated by abstract machines and refinements have been discharged,

- Assessing the mathematical correctness of the specific proof rules added to discharge proof obligations.

4.2 Expert Validation

In section we describe the procedures used in the railway and smart card case studies for reviewing B models by domain experts. The review by domain experts verifies if the model describes the requirements and physical properties properly.

The reviewing by domain experts is an important task as many errors can be rapidly discovered. And discovering errors as soon as possible is one of the aims of formal methods. Therefore, reviewing the specification and source code is mandatory for good development. However, some precautions need to be taken in order to have a good review. For example, the review must be performed by a domain expert that has not participated neither in the development of the system nor in the production of informal specification in order to avoid any bias.

4.2.1 Transportation Case Study

In the railway case study the expert validation performed several verifications through all the phases of the system model development:

- review of the semiformal properties taking in account the informal documents;

- conformity analysis between B model and semiformal properties;

- proof based upon B model;

- safety analysis based upon semiformal document.

Figure 4.1 shows all the validations done by the domain experts. The two higher validation activities, the review of the properties, events and automatons document, and the conformity analysis, are detailed below.

4.2.1.1 Validation of the Properties, Events and Automatons Document

The entire system was modelled using Event-based B. This first modelisation will be proved, as far as possible since some predicates will not be provable: these non-provable predicates will lead to discover some mistakes and/or incompleteness in the system specification and/or in the modelisation.

However, the system specification must be complete and correct enough to write a first "nearly correct" model, because if not, we may have a "completely false" model with a large number of false proof obligations, and we will not be able to correct the model by correcting all mistakes successively.

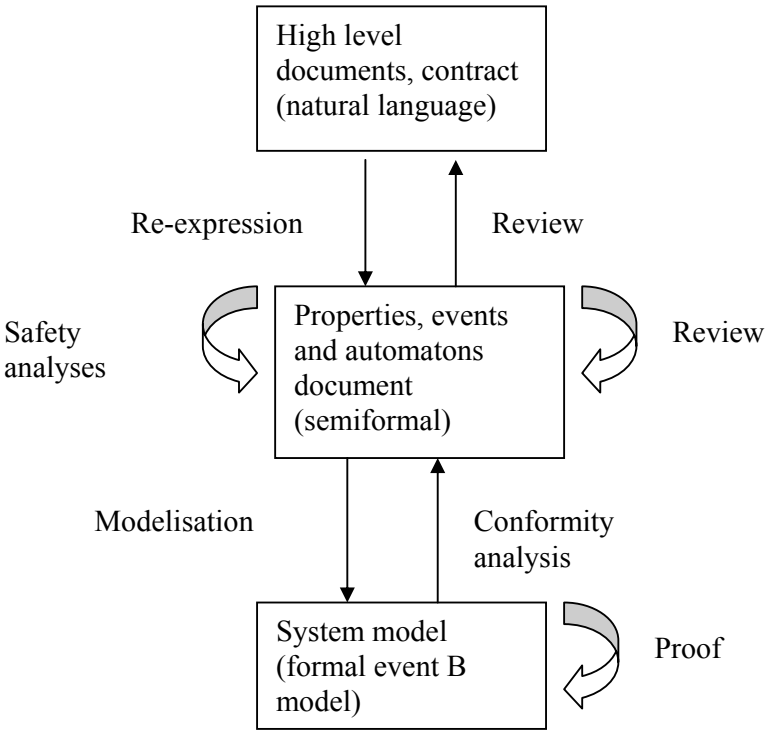


Figure 4.1: Expert validation.

This means that the system specification must be validated before the modelisation task (the proof of the system model cannot be considered as a sufficient validation of the system specification).

This document is reviewed and then approved by system experts, with two points of view: correctness of the documents regarding previous documents and safety of the system described in the properties, events and automatons document.

4.2.1.2 Completeness and Correctness of the System B Model

The system model must be strictly equivalent to the system properties, events and automatons. This means that all properties, events and automatons must be correctly described in the system model. Moreover, all pieces of information of the model must be equivalent of the connected properties, events and automatons. This task is performed by reviewing.

In the properties, events and automatons document, each property, event is associated to a unique identification number. Each automaton transition and state is identified with also with a unique identification number. This enables a rigorous traceability from the document to the B model.

The classification, with several level of abstraction, of the system properties, events and automatons cannot be validated in itself, since there is not only one solution, but an infinite number of « more or less relevant » solutions. However, the reviewer may give advice if it is obvious that some classification choices may cause difficulties in the modelisation task.

If the classification of properties, events and automatons is relevant, each level of abstraction is connected with one refinement of the system model. In that case, the rereading task is performed step by step, for each refinement (which allows detecting errors as soon as possible). The task is complete when the last refinement is revised.

4.2.2 Smart Card Case Study

The review of the smart card case study has been done by experts of the application field without any prior experience in reading B specification. They compared the informal rewriting of the specification we have done previously with the B specification. The domain experts understood rapidly the B syntax without any training. The experts wrote three reports on the flaws discovered at the most abstract level of the specification. At the time the reports were delivered some of those errors have been already discovered during the refinement and proof process. But some of the remaining undetected errors were relevant. In particular, one of those undetected errors has affected 12 instructions. It concerns an incorrect transformation of the stack by those instructions. This error was not due to the translation of informal description to the formal modal, but to the flawed informal specification written for the type verifier. So besides correcting the formal model, we also corrected the informal specification. Another error has been discovered during the review of the B model concerning a badly modelled instruction (*checkcast* instruction).

4.3 Safety Analysis

While developing safety-critical systems, it is necessary to satisfy not only functional requirements defining the set of tasks to be performed by the system, but also safety requirements describing which characteristics the system should possess in order to ensure safety.

4.3.1 Hazard Analysis

The safety requirements result from safety analysis. A paramount role in safety analysis plays hazard analysis. A *hazard* is a situation in which there is actual or potential danger to people or

to the environment the system operates in [Leveson95, Storey96]. Hazard analysis allows the designers to identify the potential danger associated with a system.

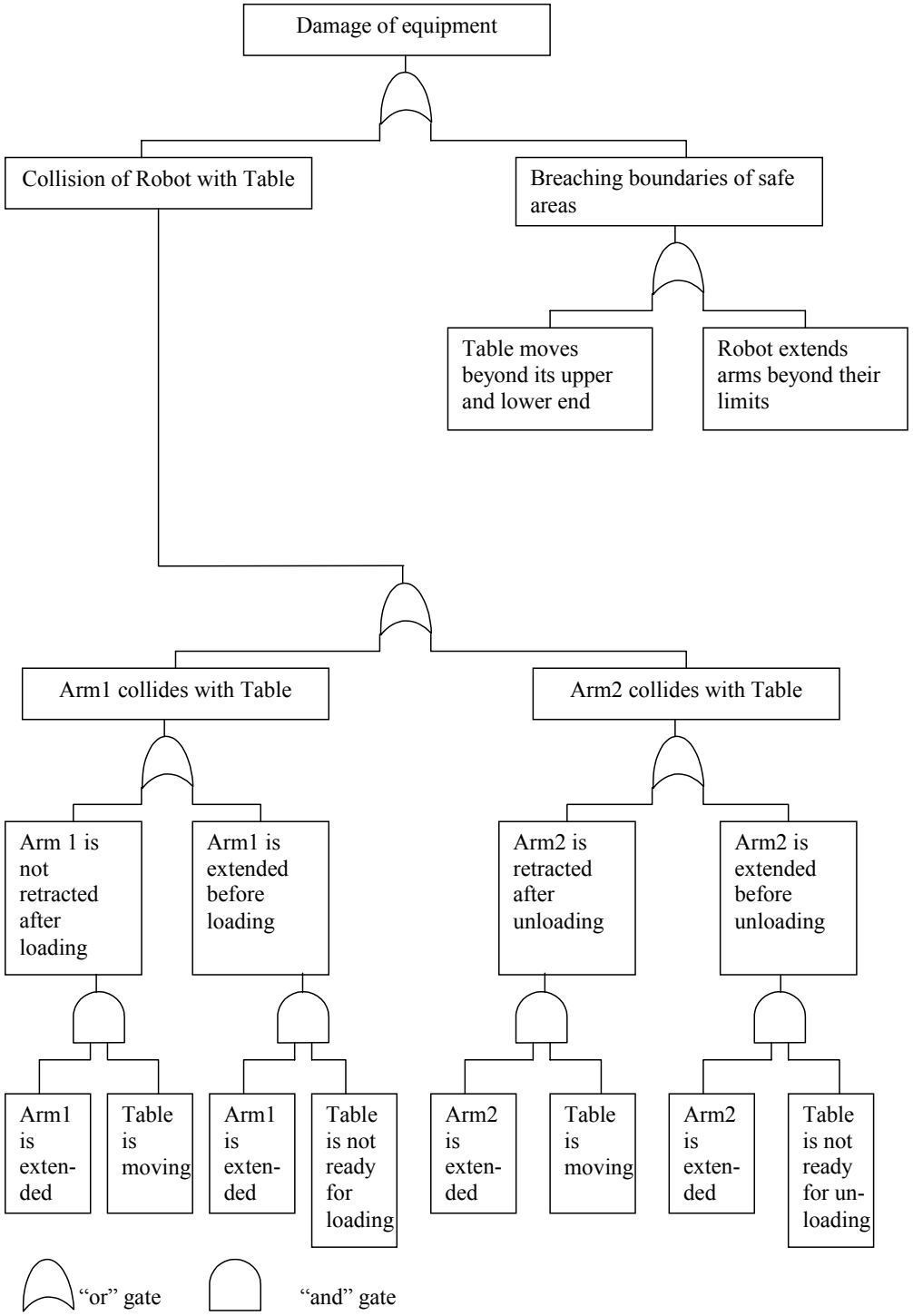


Figure 4.2: Example of Fault Tree Analysis for the robot analyser.

Hazard analysis is a set of techniques, which provide the mechanism for identification of safety requirements. Each of techniques gives a different insight into the characteristics of the system under construction. One of the most widely used techniques is Fault Tree Analysis (FTA)

[Leveson95, Storey96]. FTA is a deductive safety analysis technique. It is a top-down approach applied during the design stage. As an input FTA takes information about functions of the system and possible dangers associated with them. The result of the FTA is an identification of component faults that result in different hazardous situations. Each fault tree has a root representing a hazardous situation. The tree traces the system to the lower component level to reveal the possible causes of the hazard. Moreover, each fault tree provides a logical representation of a hazard in terms of faults of the components.

Below we illustrate how FTA allows us to formulate safety requirements, which the controlling software should satisfy. Consider the Robot Analyser described in Section 3.1.2. Preliminary hazard identification establishes that the main danger associated with the system is a damage of the equipment. This hazard forms the root of the fault tree given in Figure 4.2. Each of two events – "Collision of Robot with Table" and "Breaching boundaries of safe areas" might cause this hazard. These events are connected by disjunction, *i.e.*, logical gate "or" with the root of the fault tree. Next we analyse a logical connection the events, which in their turn lead to "Collision of Robot with Table". The analysis continues until a desired level of formulation of safety requirements is reached. For instance, safety requirements obtained using the fault tree given in Figure 4.2 are as follows:

- When the Robot is ready to load the Analyser, arm 1 may extend only if the table is in its middle position.
- When the Robot is ready to unload the Analyser, arm 2 may extend only if the operating table is in its lowest position.
- The table may move only when the respective Robot arm has retracted.
- The table must not move beyond its upper and lower end position
- The robot must not extend its arm beyond their limit positions.

4.3.2 Software Development

While designing software for safety-critical systems, it is necessary to consider the possible safety implications of the actual development, *i.e.*, to ensure that the suggested design does not introduce additional hazards. Moreover, we should ensure that controlling software reacts promptly on hazardous situations by trying to force the system back to a safe state. Clearly, this goal can be achieved only if the information provided by safety analysis is taken into account in the process of software development.

We propose an approach for conducting software development hand-in-hand with safety analysis as presented in Figure 4.3. Observe, that both the safety analysis and stepwise program development conduct reasoning from an abstract level to a concrete level. Safety analysis starts by identifying hazards that are potentially dangerous, and proceeds by producing detailed descriptions of them together with finding the means to cope with these hazards. The refinement process starts by producing an abstract specification describing what should be done, rather than how. Each refinement step allows us to incorporate more implementation details in the abstract specification, so that we end up with a program. Therefore, we can incorporate the information

that becomes available at each stage of the safety process by performing corresponding refinements of the initial specification.

We, as other researchers [Leveson96], argue that only if safety and reliability attributes are considered from the early stages of the system development, can the required dependability of the system can be achieved.

4.3.3 Dependability Impairment

A safety-critical system is typically a control system with a computer-based controller managing a plant. There are two main entities involved: an environment and a controller. While developing a specification for a control system, we model the behaviour of the overall system, that is, the physical environment and the controller together. This allows us to state explicitly the assumptions that we make about how the environment behaves. By specifying the overall system we can abstract away from means of interaction between the controller and the plant (via sensors and actuators) and, as a result, obtain a more succinct initial specification. In the later stages of the development process, when the desired level of detail has been achieved, we separate the controller and the plant, thus, obtaining the specification of the controller.

The plant behaviour evolves according to the physical processes involved and the control signals provided by the controller. The controller observes the behaviour of the plant and adjusts it to guarantee the intended plant functionality. While observing the plant behaviour the controller reacts to certain events requiring its intervention. Thus, this kind of system is a typical example of a reactive system.

The controller bases its decision on the information that it obtains from the sensors. Due to faults of the system components and imprecision in the sensors, the controller obtains an approximation of the equipment state rather than a real state. Failure occurrences deteriorate this approximation and potentially lead to dependability impairment by:

- loss of precision in the function execution, *i.e.*, a decreased reliability of the system,
- system failure as a result of violation of safety, complication in synchronizing the involved machines (which has implications for both safety and reliability).

Therefore, failure handling constitutes an important aspect of safety-critical system development.

In order to build a controller able to withstand component failures the following actions are required:

- to understand the nature of faults and their impact on the overall system,
- to decide on the detection procedures, repair procedures, functioning of the system in presence of failures, and
- to introduce an effective maintenance procedure.

These actions are performed within safety analysis. To incorporate the results of safety analysis in the refinement process the following methodological aspects should be addressed:

- modelling fault occurrence;
- modelling fault detection;
- modelling system behaviour in the presence of faults;
- modelling different operating modes and transition between modes including the system failure; and
- the design and specification of a maintenance procedure.

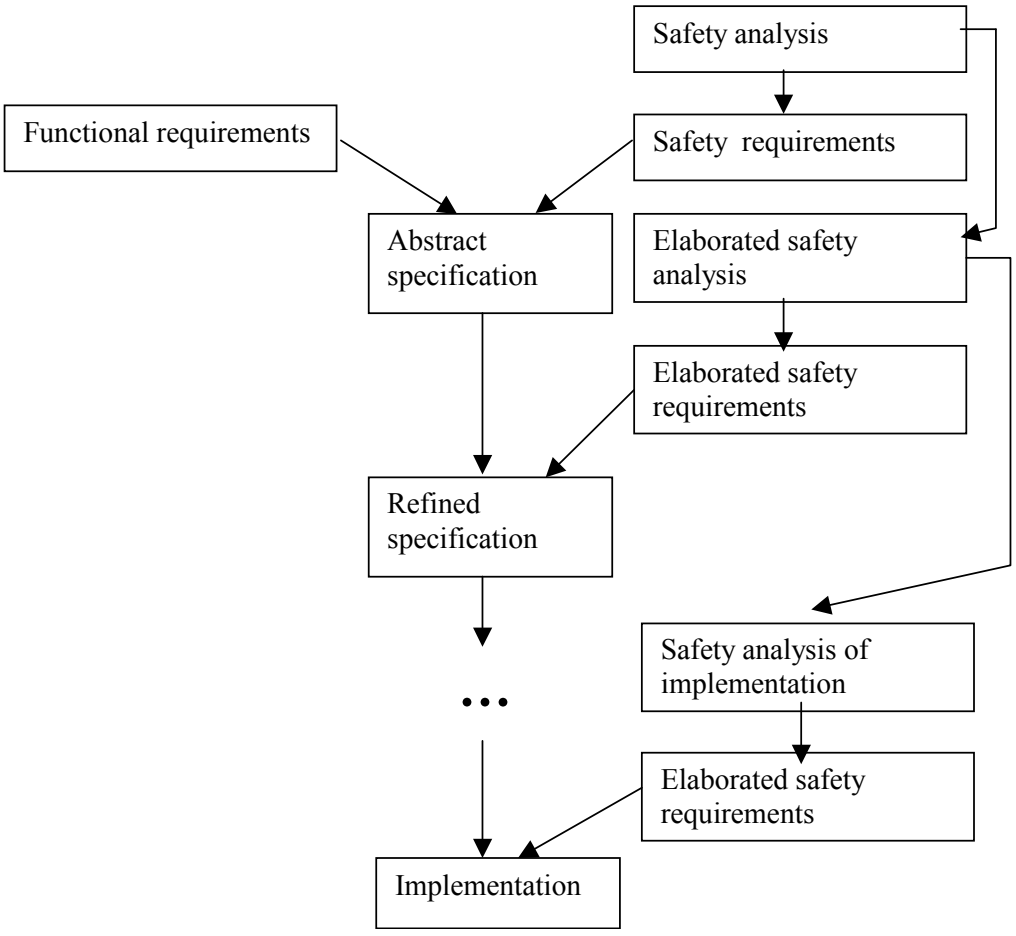


Figure 4.3: Safety analysis and stepwise system development.

4.3.4 Formal Specification and Safety Analysis

In the healthcare case study the initial specification of the system (see Figure 3.6) is rather abstract – it merely models transition between states. However, already in the initial specification we reserve a possibility of fault occurrence and system failure. The actions `Service1_fail` and `Service1_notok` model failure of execution of command.

```
Service1_fail = SELECT cmd = serv1 ∧ state = Idle THEN state := Suspension END

Service1_notok = SELECT state = Service1  THEN state := Suspension END
```

There is also possibility of spontaneous fault occurrence even while the service is not requested, as modelled by the action `Service_notready`:

```
Service_notready =  SELECT state  = Idle THEN state := Suspension END
```

In all the actions presented above, the system reacts on fault occurrence by entering the state `Suspension`. From that state the system tries to execute a recovery procedure and continue functioning as specified by the action `Remedy`:

```
Remedy = SELECT state = Suspension THEN state :∈ {Idle,Service1,Service2,Service3} END
```

We model failure of the system by action `Failure`:

```
Failure = SELECT state = Suspension THEN state := Abort END
```

This action represents transition of the system in a fail-safe state. The failure of the system occurs if the fault tolerance limit has been reached and the system cannot carry out its functions anymore.

4.4 Vulnerability Analysis

4.4.1 Smart Card

Certification Process

Application security can be enforced using rigorous processes. In order to enforce even more this security, smart-card manufacturers submit more and more often their products and processes to evaluation. The Common Criteria for Information Technology Security Evaluation (CC) standard define a set of criteria to evaluate the security properties of a product in term of confidentiality, integrity and availability. The CC focuses mainly on the first part of the lifecycle: requirements, specifications, design, development and test. The CC takes into account the security requirement documents, which are a complement of the general requirement document. Indeed, the CC are only concerned by the security aspects of the system.

The Target Of Evaluation (TOE) is the part of the product or the system that is subject to evaluation. The assurance that the security objectives are achieved is linked to:

- The confidence in the correctness of the security functions implementation, *i.e.*, the assessment whether they are correctly implemented,
- The confidence in the effectiveness of the security functions, *i.e.*, the assessment whether they actually satisfy the stated security objectives.

The Evaluation Assurance Levels (EAL1 to EAL7) form an ordered set to allow simple comparison between TOEs of the same kind. At EAL5 level the assurance is gained through a **formal** model of the TOE **security policy** and a **semiformal** presentation of the functional specification and high-level design and a **semiformal** demonstration of *correspondence* between them. Note that the analysis must include validation of the developer's covert channel analysis and a strong vulnerability analysis. The last EAL levels require a formal in-depth and exhaustive analysis.

Covert Channel Analysis

The covert channel analysis allows the existence of exploitable channels to be discovered. These covert channels can be classified into two types:

- storage channels, and
- timing channels.

Although fundamentally the same, storage channels differ in the way that information is encoded. In a storage channel, there is a shared global variable in the system that acts as the medium for information transfer. A timing channel requires the ability to reference a real time clock. If the receiver detects a timing difference in code execution, it can infer information. Therefore, information may be transferred. A covert channel can exist in the smart card if between two (or more) applets the two subject have the potential to communicate (through the shared interface) and if this communication is not allowed under the non-discretionary security policy.

For this purpose we have developed a tool and a methodology for automatically analysing this potential illegal information flow as described in [Bieber99] and [Bieber00]. The PACAP project[9] aimed to check Java Card applet interaction. More precisely, it checks the data flows between objects on the card by static analysis prior to applets downloading, for a given configuration. Information flows between applets that share data through shareable interfaces are verified. The sharable interface is the means to transfer information from one applet to another in Java Card. A shareable interface can be called by an applicative command or an external call. In all the interactions, we check if the level associated with all the variables (system and application) does not exceed the allowed sharing level that defines the security policy. The tool verifies automatically if a set of applications correctly implements a given security policy, as proposed in [Girard99].

An application is composed of a finite number of interacting applets. Each applet contains several methods. For efficiency reasons, the number of applets and methods analysed when a new applet is downloaded or when the security policy is modified are limited. The method to verify the security property on the application byte code is based on three elements:

- *Abstraction*: all values of variables are abstracted by computed levels,
- *Sufficiency*: an invariant that is a sufficient condition of the security property is verified;
- *Model checking*: this invariant is verified by model checking.

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[9] The PACAP project was partially founded by MENRT contract number 98B0252.

The abstraction mechanism and the invariant definition have been described in [Bieber00]. The tool needs as inputs, a representation of the lattice and the configuration (*i.e.,* the set of applets). With this information the tool transforms automatically the byte code into a formal semantics, adds the relevant invariant and performs the verification of the invariant.

In the example described in the previous reference, a quarter of the properties are verified within 5 seconds and 90% are verified with less than 10 minutes. One property required 23 minutes to be checked on a Sun Ultra 80. Within this tool it is possible to provide evidence to the evaluator that the applet respects the security policy. If a property cannot be validated, a counter example is given, which need to be analysed. In fact not all the counter examples express a real covert channel due to the abstraction that have been made.

This tool provides an efficient means that covers the EAL 7 requirement: "The analysis documentation shall provide evidence that the method used to identify covert channel is systematic".

Vulnerability Analysis

At the highest level (EAL7) it is not required to use formal methods. But as demonstrated in the covert channel analysis, any form of automation saves time. One of the requirements is to provide a justification that the analysis completely addresses the TOE deliverables. The evaluator will use the analysis that the vulnerabilities have been addressed but will also perform an independent vulnerability analysis. Vulnerability analysis deals with the threats that a user will be able to discover flaws that allow unauthorised access to resources (*e.g.*, data). Such an attack allows the authorised capabilities of other users to be interfered with. The vulnerability analysis enables checks to be made on the coherence of the security policy and the enforcement of this security policy in the product implementation. It requires first an informal model of the interactions that take place between applets and between the selected applet and the terminal. The vulnerability analysis consists in examining all the application responses. The set of this responses are called traces. The analysis of these traces determines whether data is accessible or not. To perform this analysis, we have to model inference rules that represent dependence and the security policy. The FDR model checker [FDR97] is used to check properties over the model by examining the set of traces, looking for illegal transactions or private data that are publicly accessible. If data is accessible, we determine if it is in accordance with the security policy.

4.5 Proof

This section shows how to use Atelier-B to validate a B model; it describes strategies for automated proof, interactive proof, and rule management.

4.5.1 An Automated Proof Strategy

The facility to perform proofs within the B formalism enables direct and powerful validation. A functional property that a system should possess can be demonstrated using proof. The disadvantage of performing proofs is the amount of effort it general requires.

4.5.1.1 The General Approach

When developing in B, starting interactive proof too early may have dramatic consequences in term of cost and delay, because models are likely to evolve and proof work would be lost. On the other hand, if no proof is undertaken in the early stages of the development, high level specification may be wrong and, again, models would be corrected. A maxim is:

Do not start interactive formal proof too soon, but start the proof phase before completing all the models

One good way[10] to proceed is to mix automatic proof and visual inspection of the Proof Obligations (PO). The objective is to achieve a sensible balance between the confidence that no false PO remains, and the time required to achieve this confidence.

The general approach is follows the flowchart of Figure 4.4:

- Write abstract machine according to requirements document,
- Control the correctness of the formal expression of the needs,
- Start automatic proof on this abstract machine,
- If some proof obligations are not automatically demonstrated, check them quickly by reading them and verify that they seem to be true. If any are false, the abstract machine should be corrected.
- Write implementation component,
- Compare implementation versus abstract machine,
- Start automatic proof on this implementation,
- If some proof obligations are not automatically demonstrated, check them quickly by reading them and verify that they seem to be true. If some of them are false, the abstract machine or the implementation should be corrected.
- Make the formal demonstration of the remaining PO in the abstract machine and in the implementation, using interactive prover.

[10] It has proved to be efficient on many industrial-size projects.

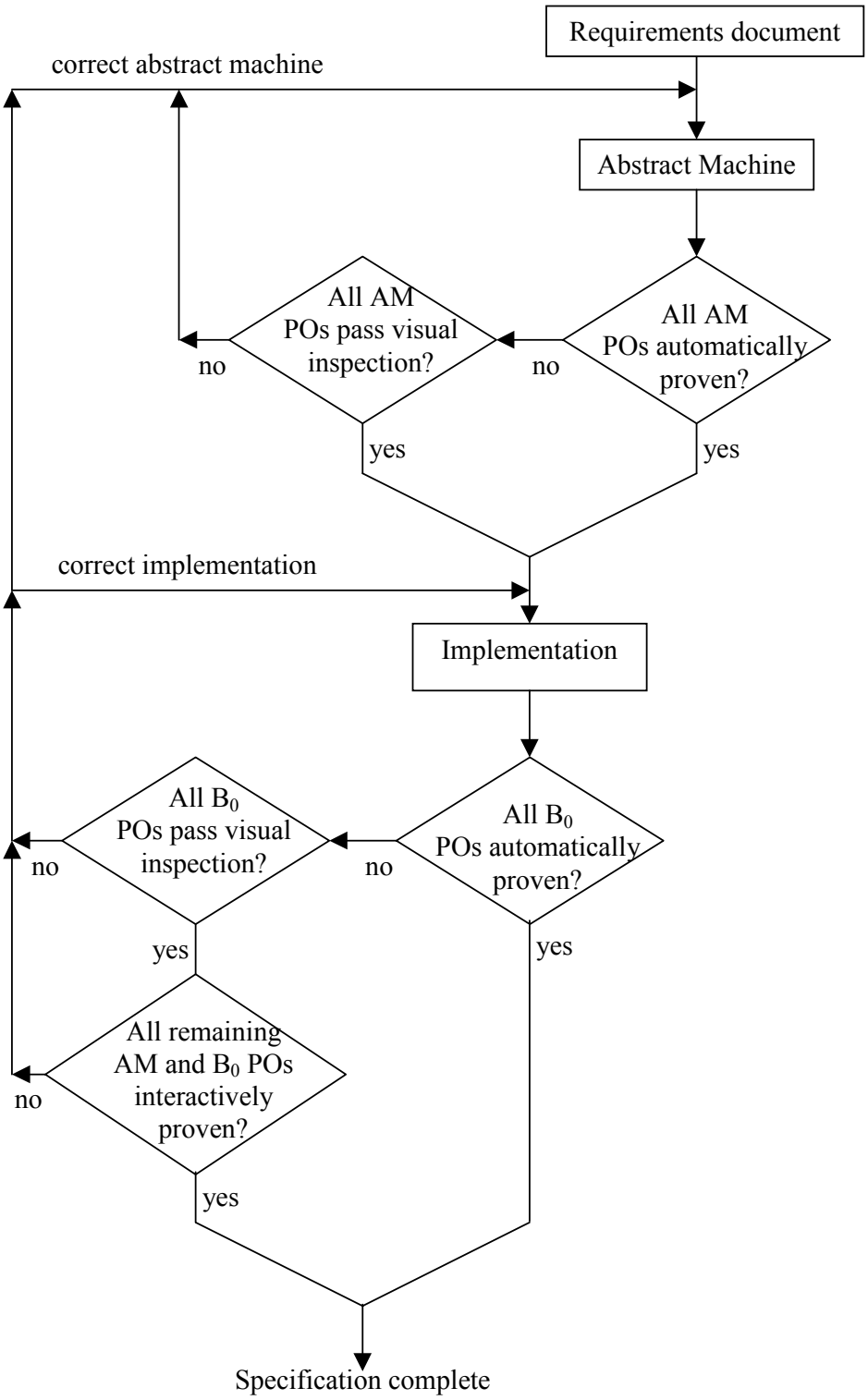


Figure 4.4: Automated proof strategy.

4.5.1.2 Automatic Proof Rate as a Quality Indicator

Automatic proof rate is a good indicator of the suitability of a B model. Most of the time, if the proof rate is too low, the B model needs to be reshaped. From our experience, large-scale projects have an automatic proof rate between 70% to 95%. Locally, some components may have a low rate (a small percentage automatically proven), depending on their complexity and the operations that are used to construct the B model.

Reducing model complexity

A model may be too complex, because:

- Too many details are included in the top-level specification components; this may be overcome by:
  - identifying and using only those variables that need to be present at this level; and
  - introducing details in the lower components.
- Too heavy use has been made of sequencing; this may be overcome by:
  - decomposing operations (the `LOCAL_OPERATIONS` clause may be helpful).

Constructing models that can be easily manipulated by the automated prover

Some operations may confuse the automatic prover and lead the user to perform interactive proofs. These operations are:

`closure`, `union`, `inter`.

There are two main reasons for this:

- These operations are difficult to manipulate in a general way. Many proofs need to add very particular rules, that are likely to be difficult to establish and demonstrate, which have little chance to be reused for other projects.
- The prover rules database has been built mainly from railway developments which have not used these operations. So there are none or few rules regarding the manipulation of these operations.

There are three general techniques for dealing with this problem.

- Use quantified predicates instead of these operations.
- Use specifically created rewriting rules to rewrite these operations into more manageable operations, expressions or predicates (if applicable).

For other operations, particular rules should be added for arithmetic operators as general sum, general product, and modulo.

In Atelier-B 3.6, a new package of formally validated rules will be available, providing axiom-like rules related to these symbols arithmetic operators that were badly manipulated by older versions of the prover.

Finally, if the extended prover rules database still does not contain required rules, to simplify expressions or predicates containing these operations, .they can be manually added.

4.5.1.3 Playing with Forces

The automatic prover is composed of:

- a **rule database**, containing about 2800 mathematical rules. These rules are grouped in packets called *theories* and can each *theory* be activated independently from the others.
- a **proof engine**, whose rôle is to:
- process hypotheses (simplification, stack loading),
- simplify the goal, by
- applying heuristics. Those heuristics are general (deduction, skolem transformation, …) and don't address very specific goal patterns.
- using some rule *theorie*s from the rule database. Rule theories are triggered by the proof engine if the rule contained in them match with the current goal.

The efficiency of the automatic prover can be parameterised by using different forces. Forces are indexed according to the complexity of the mechanisms involved in the proof process (see Table 3).

Force	Proof time per lemma	Performance
0	Less than 10 seconds	70 %
1	From some seconds to 3 minutes	+ 1 %
2	From some minutes to tens of minutes	+ 3 %
3	From tens of minutes to hours	+ 1 %
Fast	Less than 3 seconds	30 %

Table 3: Efficiency of the automatic prover for different forces.

Force 0 is the most efficient force: its original specification is to have a maximum proof duration of 10 s. Hypotheses are processed differently from higher forces: each hypothesis is simplified, by using a dedicated part of the rule database only used for that goal, then loaded in the stack. The proof engine then try to apply rules from rule theories that can match with the current goal. In case of success, the proof engine stars a new iteration. In case of failure, some heuristics are tried (skolem transform, simplification of negated predicates, simplification of the last

hypotheses loaded in the stack, …). In case of failure (nothing new has been obtained), the prover stops. Otherwise, the proof engine starts a new iteration.

Forces 1, 2 and 3 do not process hypotheses the same way: the proof engine tries to demonstrate each hypothesis. If it succeeds, the hypothesis is obviously redundant with hypotheses previously loaded and therefore, is deleted. In case of failure, the simplified hypothesis (the proof engine always try to simplify incoming goals) is loaded in the stack. That way, the stack is supposed to contain a minimal set of hypotheses. The drawback of this approach is the time required to process each hypothesis: about 1 minute for 100 hypotheses. The heuristics used by the proof engine are also different: some of them introduced in forces 1, 2 and 3 allow the proof engine to try to demonstrate missing information that would "probably" help proof (for example, a missing hypothesis that would enable to fire a rule related to the goal). This approach is obviously time-consuming and should never be used when trying to prove a B component for the first time.

4.5.2 Manual Proof

There is one major rule related to manual proof:

Inspect PO before performing any manual proof activity.

4.5.2.1 PO Inspection

PO inspection can be performed either with the PO viewer or the interactive prover.

Difficult POs should be evaluated first. So inspection should begin from the end of the list of all unproved POs, as POs are ordered by their degree of difficulty starting with the simplest.

PO inspection consists of five steps:

- goal reading (interpret and isolate verified constraint)
- justification (use the physical meaning of the component),
- key hypothesis selection (search matching hypothesis),
- intuitive demonstration (reuse justification and examine used rules),
- notes and trials (write down used simplifications and try a quick demonstration).

Quick demonstrations should be preformed within a time limit (no more than 10 minutes for example). First try predicate prover. Do not use more than 5 commands. Try to generalise a demonstration to other PO.

If a false PO is found, the B model should be corrected before continuing. It should be verified that this PO really is false, as this may not be obvious: a good method is to exhibit a counter-example. Introducing this counter-example in the B model allows the source of errors to be localised.

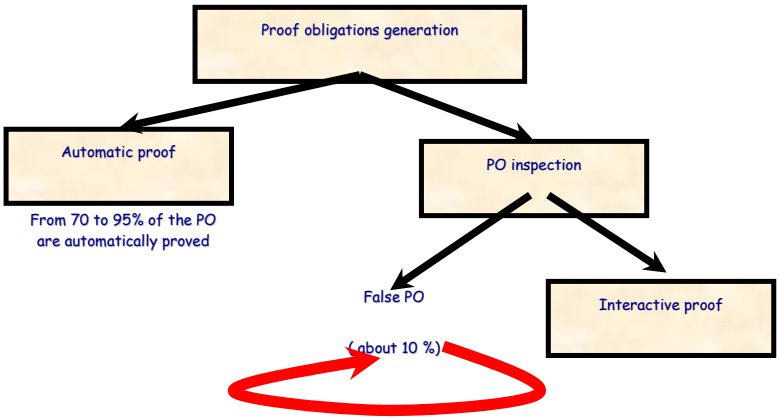


Figure 4.5: Proof rate for the automatic prover.

4.5.2.2 Proving Manually

Proving is part of the development process, just as much as the creation of models. Indeed, having written a model will help the user to prove and modify it. In a typical project, the automatic prover helps to locate errors in the B model, which represent about 10% of the total POs. Another reason is that model complexity is tightly coupled to proof complexity: two models may be semantically equivalent, but their formulations (in particular the operations that are used), may mean that one is easier to prove. Efficiency in a B development is achieved through the construction of easily (most of the time) proved models.

The average proof rate that has been measured is between 10 and 20 POs per day, including debugging, for the lifetime of the project.

We can consider the following rules as good axioms:

- Use no more than 10 steps in an interactive demonstration. (There are exceptions; the maximum known is about 550 steps.)
- Add a user rule if the situation is not covered well by the prover rule database or proof mechanisms.
- If a predicate is demonstrated for more than one PO, add an assertion in the B model.

A new assertion should be added after all the other assertions. If this is done, the proof status is not modified and demonstrations are not lost. Indeed, when a model is modified, the merger (part of the PO generator) intelligently updates the proof status.

If previously, a PO was $H_1 \wedge H_2 \Rightarrow G$, the proof status is:

- kept if the new PO is $H_1 \wedge H_2 \wedge H_3 \Rightarrow G$ (because it can be logically deduced),
- initialised if the new PO is $H_1 \wedge H_3 \Rightarrow G$.

Similarly, when an invariant is corrected, by adding a variable or a property, the new predicate should be added at the end of the clause. This limits the risk of loosing interactive demonstrations, since the proof tree has been modified and it is unlikely that the demonstration will need to be replayed.

4.5.2.3 Example of Manual Proof

To illustrate the use of interactive prover we will describe the steps necessary to prove the POs of the RailwayT machine presented in Figure 2.6 and Figure 2.7.

We start by presenting the `Occupy` operation, omitted in Figure 2.7 for simplicity. This operation describes a train `t1` entering section `c1`. The first `LET` expression declares the local variable `newocc` that holds the set of sections currently occupied by trains. The variables `occ`, `pos`, and `atime` are updated to reflect train `t1` entering a new section. The second `LET` declares two local variables, `newres` and `diff`: variable `newres` holds the new restrictive sections; and `diff` holds the sections that have become restrictive with the occurrence of `Occupy`. The variable `rtime` (that describes the time at which a section goes restrictive) is updated for the new restrictive sections.

```
Occupy(t1,c1) =
    PRE
        t1 ∈ TRAIN ∧ c1 ( SECTION
    THEN
        LET newocc BE newocc = ran(pos  {t1  c1})
        IN
            occ := newocc  ||
            pos(t1) := c1  ||
            atime(t1) := cur  ||
            LET      newres, diff BE  /* Several sections may become restrictive: */
                newres = { cc | cc ( SECTION  (  Restrictive(cc,newocc) } (
                diff = newres-res
            IN
                rtime := rtime  (diff(({cur})  ||
                res:=newres
            END
        END
    END
```

After applying the automatic prover to the RailwayT machine 41 POs are generated, but only 2 POs remained unproved. These two POs are related to the Occupy operation, and they were proved using the interactive prover of Atelier-B (also called manual prover). In here, we will describe in detail the interactive proof of the following PO:

```
(atime<+{t1|->cur})(tt)+XX<=cur &
"`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res & atime(tt)>=rtime(pos(tt)) &
atime(tt)+XX<=cur => tt: braking)) is preserved by the operation - ref 3.4'" &
=>
tt: braking
```

This PO states that if a train tt is in a restrictive section, and the train has arrived in that section after the section went restrictive, then it must be braking. This PO verifies if the Occupy operation preserves the second case of the braking property:

```
BrakingProperty(i) =
…
∀tt.( tt ∈ TRAIN ∧ pos(tt) ∈ res ∧
              atime(tt) ≥ rtime(pos(tt)) ∧ atime(pos(tt)) + XX ≤ i
       ⇒ tt ∈ braking )
```

We start the by splitting the proof in two cases. First, we consider the case where `t1 = tt`, *i.e.*, `tt` is the train that has just entered section `c1` by invocation of `Occupy(t1,c1)`. After that we start the deduction by invoking the command **`dd`**.

```
PRI > dc(t1 = tt)
Starting Do Cases
        t1 = tt => ("`Local hypotheses'" & newocc = ran(pos<+{t1|->c1}) &
        newres = {cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
        diff = newres-res & tt: TRAIN & (pos<+{t1|->c1})(tt): newres &
        (rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) &
        (atime<+{t1|->cur})(tt)+XX<=cur &
        "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
        atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking))
        is preserved by the operation - ref 3.4'"
        =>
        tt: braking)

PRI > dd
Starting Deduction
        "`Local hypotheses'" &
        newocc = ran(pos<+{t1|->c1}) &
        newres = {cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
        diff = newres-res &
        tt: TRAIN &
        (pos<+{t1|->c1})(tt): newres &
        (rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) &
        (atime<+{t1|->cur})(tt)+XX<=cur &
        "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking)) is preserved by the
operation - ref 3.4'" &
        =>
        tt: braking
```

Next, as we are dealing with the case where `t1 = tt`, the variable `tt` can be replaced by `t1`, which is done by the command **`eh`**. After this substitution the consequent of the goal becomes `t1: braking`, and a set simplification is applied to the goal (command **`ss`**).

```
PRI > eh(tt,t1,Goal)
Starting use Equality in Hypothesis
        "`Local hypotheses'" &
        newocc = ran(pos<+{t1|->c1}) &
        newres = {cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
        diff = newres-res &
        t1: TRAIN &
        (pos<+{t1|->c1})(t1): newres &
        (rtime<+diff*{cur})((pos<+{t1|->c1})(t1))<=(atime<+{t1|->cur})(t1) &
        (atime<+{t1|->cur})(t1)+XX<=cur &
        "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
                atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking))
                is preserved by the operation - ref 3.4'" &
        =>
        t1: braking

PRI > ss
Begin SimplifySet
        "`Local hypotheses'" &
        newocc = ran({t1}<<|pos)\/{c1} &
        newres = {cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
```

```
        diff = newres-res &
        t1: TRAIN &
        c1: newres &
        (rtime<+diff*{cur})(c1)<=cur &
        XX<=0 &
        "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
                atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking))
                is preserved by the operation - ref 3.4'" &
        =>
        t1: braking
```

We start a deduction by calling **dd**, which adds the goal antecedents as new hypotheses. To prove the new goal we add two hypotheses: the first one was already in the antecedent of the goal above; the second one is part of the definition of constant XX. After calling the auto-prover the proof for the case where `t1 = tt` is concluded, and now we have to prove the other case where $t1 \neq tt$.

```
PRI > dd
Starting Deduction
        t1: braking

PRI > ah(XX<=0)
Starting Add Hypothesis
        XX<=0 => t1: braking

PRI > ah(XX>=1)
Starting Add Hypothesis
        1<=XX => (XX<=0 => t1: braking)

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        not(t1 = tt) => ("`Local hypotheses'" & newocc = ran(pos<+{t1|->c1}) &
             newres ={cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
             diff = newres-res & tt: TRAIN & (pos<+{t1|->c1})(tt): newres &
             (rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) &
             (atime<+{t1|->cur})(tt)+XX<=cur &
             "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
             atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking))
             is preserved by the operation - ref 3.4'"
             =>
             tt: braking)
```

To add the antecedents to the hypotheses we invoke the twice the command dd (this is necessary, because the goal has the form $A \Rightarrow A' \Rightarrow C$), the first invocation just adds the antecedent `not(t1 = tt)`, while the second invocation adds the remaining antecedents.

```
PRI > dd
Starting Deduction
        "`Local hypotheses'" &
        newocc = ran(pos<+{t1|->c1}) &
        newres = {cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
        diff = newres-res &
        tt: TRAIN &
        (pos<+{t1|->c1})(tt): newres &
        (rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) &
        (atime<+{t1|->cur})(tt)+XX<=cur &
        "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
        atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking))
        is preserved by the operation - ref 3.4'" &
        =>
        tt: braking

PRI > dd
Starting Deduction
        tt: braking
```

Now the current goal is `tt: braking`. To prove that goal several hypotheses are added. These three hypotheses were already given in the antecedent of the implication above.

```
PRI > ah((pos<+{t1|->c1})(tt): newres)
```

```
Starting Add Hypothesis
        (pos<+{t1|->c1})(tt): newres => tt: braking

PRI > ah((rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt))
Starting Add Hypothesis
        (rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) =>
((pos<+{t1|->c1})(tt): newres => tt: braking)

PRI > ah((atime<+{t1|->cur})(tt)+XX<=cur)
Starting Add Hypothesis
        (atime<+{t1|->cur})(tt)+XX<=cur => ((rtime<+diff*{cur})((pos<+{t1|-
>c1})(tt))<=(atime<+{t1|->cur})(tt) => ((pos<+{t1|->c1})(tt): newres => tt: braking))
```

As we know that `tt ≠ t1`, all the hypotheses we just added bcould be simplified. For instance, the expression `((pos<+{t1|->c1})(tt): newres` can be simplified to `pos(tt): newres`, which implies that `tt` is a restrictive section. To be able to do those simplifications, first we have to prove that `tt` (before the occurrence of `Occupy`) was already defined for both functions `pos` and `atime` (the time a train has entered its current position).

```
PRI > ah(tt: dom(pos))
Starting Add Hypothesis
        tt: dom(pos)

PRI > pr
Starting Prover Call
        tt: dom(pos) =>
         ((atime<+{t1|->cur})(tt)+XX<=cur =>
         ((rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) =>
         ((pos<+{t1|->c1})(tt): newres => tt: braking)))

PRI > ah(tt: dom(atime))
Starting Add Hypothesis
        tt: dom(atime)

PRI > pr
Starting Prover Call
        tt: dom(atime) =>
         (tt: dom(pos) => ((atime<+{t1|->cur})(tt)+XX<=cur =>
         ((rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) =>
         ((pos<+{t1|->c1})(tt): newres => tt: braking))))

PRI > dd
Starting Deduction
        tt: dom(pos) => ((atime<+{t1|->cur})(tt)+XX<=cur =>
         ((rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) =>
         ((pos<+{t1|->c1})(tt): newres => tt: braking)))

PRI > dd
Starting Deduction
        (atime<+{t1|->cur})(tt)+XX<=cur =>
         ((rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) =>
         ((pos<+{t1|->c1})(tt): newres => tt: braking))
```

We have now proved the two above hypotheses, so we can simplify our goal by applying directly rule `SimplifyRelFonXY.9`. This rule states that the substitution `(f<+{a|->b})(c) == f(c)` can be applied if `c ≠ a` and `c ∈ dom(f)`.

```
PRI > ar(SimplifyRelFonXY.9,Goal)
Starting Apply Rule
        atime(tt)+XX<=cur => ((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres
=> tt: braking))
```

Looking at our current goal we can see that the expression `((rtime<+diff*{cur})(pos(tt))` is still too complex to be dealt by the prover. So we are going to divide the current goal in two cases, depending on whether the train `tt` was already in a restrictive section or not. We start with the case where the section that `tt` is occupying has just become restrictive (`dc(pos(tt): diff)`). After having proved that `pos(tt): dom(diff*{cur})` we can then simplify the goal

using rule `SimplifyRelFonXY.8`. This rule states that `(r<+s)(a) == s(a)`, provided `a ∈ dom(s)`.

```
PRI > dc(pos(tt): diff)
Starting Do Cases
        pos(tt): diff => (atime(tt)+XX<=cur =>
         ((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres => tt: braking)))

PRI > dd
Starting Deduction
        atime(tt)+XX<=cur =>
         ((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres => tt: braking))

PRI > ah(pos(tt): dom(diff*{cur}))
Starting Add Hypothesis
        pos(tt): dom(diff*{cur})

PRI > ss
Begin SimplifySet
        pos(tt): diff

PRI > pr
Starting Prover Call
        pos(tt): dom(diff*{cur}) => (atime(tt)+XX<=cur =>
         ((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres => tt: braking)))

PRI > dd
Starting Deduction
        atime(tt)+XX<=cur => ((rtime<+diff*{cur})(pos(tt))<=atime(tt) =>
         (pos(tt): newres => tt: braking))

PRI > ar(SimplifyRelFonXY.8,Goal)
Starting Apply Rule
        atime(tt)+XX<=cur =>
         ((diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres => tt: braking))
```

After simplifying the set expression we obtain an implication with several antecedents. We can see that the conjunction of the inequalities `0<=cur-atime(tt)-XX` and `cur<=atime(tt)` is false, if `XX` is a number greater than 1.

```
PRI > ss
Begin SimplifySet
        0<=cur-atime(tt)-XX &
        cur<=atime(tt) &
        pos(tt): newres &
        =>
        tt: braking

PRI > ah(XX>=1)
Starting Add Hypothesis
        1<=XX => (0<=cur-atime(tt)-XX & cur<=atime(tt) & pos(tt): newres => tt: braking)

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        not(pos(tt): diff) => (atime(tt)+XX<=cur =>
((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres => tt: braking)))
```

Now we still have to prove the case where train `tt` was already in a restricted section before the occurrence of `Occupy`, *i.e.*, `not(pos(tt): diff)`. Again, the aim of the next commands is to simplify the expression `((rtime<+diff*{cur})(pos(tt))`.

```
PRI > dd
Starting Deduction
        atime(tt)+XX<=cur => ((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres
=> tt: braking))

PRI > ah(pos(tt)/:dom(diff*{cur}))
Starting Add Hypothesis
        not(pos(tt): dom(diff*{cur}))

PRI > ss
```

```
Begin SimplifySet
        not(pos(tt): diff)

PRI > pr
Starting Prover Call
        not(pos(tt): dom(diff*{cur})) =>
         (atime(tt)+XX<=cur => ((rtime<+diff*{cur})(pos(tt))<=atime(tt) =>
         (pos(tt): newres => tt: braking)))

PRI > dd
Starting Deduction
        atime(tt)+XX<=cur =>
         ((rtime<+diff*{cur})(pos(tt))<=atime(tt) => (pos(tt): newres => tt: braking))

PRI > ss
Begin SimplifySet
        0<=cur-atime(tt)-XX &
        rtime(pos(tt))<=atime(tt) &
        pos(tt): newres &
        =>
        tt: braking
```

Because the above goal is very similar to the second case of the invariant property `BrakingProperty`, we will use the command **`ph`** to particularise that property to train `tt`.

```
PRI > dd
Starting Deduction
        tt: braking

PRI > ph(tt,!tt.(tt: TRAIN & pos(tt): res & rtime(pos(tt))<=atime(tt) & (0<=cur-
atime(tt)-XX & atime(tt)+XX<=cur) => tt: braking))
Starting Particularize Hypothesis
        tt: TRAIN
```

In the next steps we prove that the train `tt` satisfies all the antecedents of the implication within `BrakingProperty.`

```
PRI > pr
Starting Prover Call
        pos(tt): res

PRI > ah(pos(tt): newres)
Starting Add Hypothesis
        pos(tt): newres => pos(tt): res

PRI > ah(not(pos(tt): diff))
Starting Add Hypothesis
        not(pos(tt): diff) => (pos(tt): newres => pos(tt): res)

PRI > eh(diff,newres-res,Goal)
Starting use Equality in Hypothesis
        not(pos(tt): newres-res) => (pos(tt): newres => pos(tt): res)

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        rtime(pos(tt))<=atime(tt)

PRI > pr
Starting Prover Call
        0<=cur-atime(tt)-XX

PRI > pr
Starting Prover Call
        atime(tt)+XX<=cur

PRI > pr
Starting Prover Call
        tt: braking => tt: braking
```

The above goal is trivially proved by the auto-prover, which concludes the interactive proof for of PO Occupy.10.

```
PRI > pr
```

```
Starting Prover Call
        "`Local hypotheses'" &
        newocc = ran(pos<+{t1|->c1}) &
        newres = {cc | cc: CDV & not(block~[{next_block(block(cc))}]/\newocc = {})} &
        diff = newres-res &
        tt: TRAIN &
        (pos<+{t1|->c1})(tt): newres &
        (rtime<+diff*{cur})((pos<+{t1|->c1})(tt))<=(atime<+{t1|->cur})(tt) &
        (atime<+{t1|->cur})(tt)+XX<=cur &
        "`Check that the invariant (!tt.(tt: TRAIN & pos(tt): res &
        atime(tt)>=rtime(pos(tt)) & atime(tt)+XX<=cur => tt: braking))
        is preserved by the operation - ref 3.4'" &
        =>
        tt: braking
```

4.5.3 Added Rules Management

Although the prover has become more powerful, the rules database has been extended and proof mechanisms have been improved, but in some situations, rules addition is still necessary because:

- the prover cannot do anything interesting with the goal,
- using rules may save hours spent in performing interactive proofs.

Rules can be added to be used by:

- all the components of the project, by locating them in the `PatchProver` file, or
- just one component, by adding them to the associated `pmm` file.

Rules are expressed in the theory language (see Mathematical Rule Writing Guide [ClearSyD] for more details).

With the exception of *Rewriting* rules, for which class can be easily induced, there is no difference between *Backward* and *Forward* rules. So in order to keep track of their type and use, theory names should be non-ambiguous and the prover trace system should be used. This last feature allows the application of rules to be accurately traced.

All the rules that are added should be demonstrated in order to ensure that a project really is fully proven. The *Rules Proof Tools* should be used for this task. It performs the following controls:

- syntax checking,
- type-checking,
- jokers instantiation correctness,
- guards translation,
- rules translation (depending on their class: *Backward, Forward, Rewriting*), and
- demonstration.

In addition, some verification should be done by hand for all rules:

- correct calls to prover mechanisms,
- correct use of comma/maplet,
- well-definedness.

Some metrics regarding rules validation are:

- automatic rule validation (using *Rules Proof Tools*): from 30 to 85%,
- manual rule validation: approximately 20 rules per day,
- manual rule verification: approximately 50 rules per day.

Projects tend to use less user rules than before, thanks to new commands, mechanisms and extensions of the prover.

For example, the Automatic Train Protection software used more than 1200 rules in 1994, but only 22 rules in 2000.

5 Refining Models

The MATISSE project follows a strategy based on principles of *top-down stepwise tree-structured refinements and distributed design*. Top-down implies that one starts with an abstract description – the architectural model – and will end up through refinement of abstract specifications and proved transformations, with a correct system implementation. This is to be contrasted with bottom-up design, when one conceptually starts with a concrete realisation and by constructing higher-level functionality from building blocks tries to achieve and *a posteriori* prove properties of an abstract behaviour specified in the architecture. The methods used in this project belong to the former approach.

Stepwise refinements means that there are many proved intermediate design steps in the process of deriving a correct implementation for a system starting from its architectural specifications. The design step, or *transformation i* results in a new level of design $i + 1$ more concrete (less non-deterministic in terms of the B Method) than the higher level(s). For instance, a given design step may concern the behaviour of just a single component of the system. At the very end, results a series of designs for each component of a system starting with its specifications and ending with one of its implementations. The complete system design will be composed of a set of related and interconnected designs for all (concurrent) components of a whole system.

The tree-structure results because in every transformation step one has the choice (non-deterministic behaviour) between several different lower-level designs corresponding to possible implementation choices. In particular, each transformation step can lead to several concurrent designs at lower-level which can include distribution or locality dependent properties (communication/synchronisation constraints), or environment generated events (*e.g.*, reactive or timed constraints) for the implementation.

5.1 Introduction

It would be impossible to formalise a complete real system by means of a single model, because the state would be far too complicated and the number of events far too large. In order to master the modelling process, we use two complementary techniques:

- refinement, and
- decomposition.

Refining a model consists in refining its state and its events. A concrete model (with respect to a more abstract one) has a state that should be related to that of the abstraction through a, *gluing* invariant. Sometimes the concrete state is a simple extension of the abstract one so that the abstraction relation is then just a projection. But in general the abstraction relation can be any relation that is defined on all the concrete states.

5.1.1 B Refinement

An abstract machine is refined by applying the standard technique of data refinement to its state: an abstraction invariant is used to relate the state variables of the abstract system to those of the refined system and data refinement should hold between correspondingly-named operations in the abstract and refined systems (see, *e.g.*, [Abrial96,Butler96]). If `S` is a statement that acts on

abstract variables, T is a statement that acts on concrete variables, and AI is an abstraction invariant, then we write

$$S \;\; {}_{AI} \;\; T$$

for S *is data-refined by* T *under abstraction invariant* AI. The formulation of $S \;\; {}_{AI} \;\; T$ is omitted here. Details may be found in [Abrial96].

A system M is refined by system N under abstraction invariant AI if N has an operation N.a corresponding to each operation M.a, and for each such a:

$$M.a \;\; {}_{AI} \;\; N.a.$$

As well as having an operation corresponding to each of the abstract operations, a refined system may also have some auxiliary *internal* operations. These are operations that can change the state of the refined machine but whose occurrence is not observable at the more abstract level. Viewed in terms of the abstract state, such operations should have no effect, that is, they should be data-refinements of the skip statement on the abstract state. Each auxiliary operation N.h of refined machine N must satisfy:

$$skip \;\; {}_{AI} \;\; N.a.$$

The tool support available for B can be used to verify refinement between two models. Proof obligations are generated based on the abstraction invariant and these are discharged automatically and semi-automatically.

5.1.1.1 Example

To illustrate the B notion of system refinement we will use the railway example presented in Section 2.1.3.1, which is a simplification of the transportation case study.

First Refinement

The first refinement of machine Railway (see Figure 2.4 and Figure 2.5) is described in Figure 5.1. The Check operation of the system-level model (Figure 2.5) reads the values of next and occp directly to determine whether to apply the brakes. An individual train cannot read these values directly, but instead receives signals from the trackside controllers. Our first refinement introduces this signalling by introducing an auxiliary operation SendTrainMsg. The variables of the refined model are all of the variables of the abstract model plus an extra variable tmsgs mapping each train to a boolean signal being sent from a trackside controller to the train. $BOOL_{\perp}$ represents BOOL augmented with the value $\perp$, where $\perp$ models the absence of a signal. When sending a signal to train t, the SendTrainMsg operation sets the value of the signal to TRUE if it is ok for the train to continue into the next section and FALSE otherwise. The refined Check operation (also shown in Figure 5.1) reads the signal when it is available and sets the brake accordingly. The abstraction invariant used to verify that this model is a valid refinement of the original model is shown in the invariant clause of the Railway1 refinement.

```
REFINEMENT  Railway1  REFINES  Railway
INVARIANT
```

```
    tmsg ⊆ TRAIN → BOOL⊥ ∧
    ∀t.(t ∈ TRAIN ∧ tmsg(t) ≠ ⊥   ⇒
        tmsg(t) = bool(front(t) ∈ dom(next) ∧ next(front(t) ∉ ran(occp))

OPERATIONS
SendTrainMsg(t) =
    PRE t ∈ TRAIN
    THEN
        SELECT tmsg(t) = ⊥
        THEN
            tmsg(t) := bool(front(t) ∈ dom(next) ∧ next(front(t) ∉ ran(occp))
        END
    END;

Check(t) =
    PRE t ∈ TRAIN
    THEN
        SELECT tmsg(t) ≠ ⊥
        THEN
            braking(t) := ¬tmsg(t) || tmsg(t) := ⊥ || checked(t) := TRUE
        END
    END;
```

Figure 5.1: First refinement of `Railway`.

Decomposition to Parallel Sub-Systems

[Butler97] describes a technique for composing B machines in a way that corresponds to parallel composition. The interaction between the parallel machines is based on synchronisation (with value-passing) between corresponding operations of the respective machines. There are no state variables shared by the machines. Here, we apply this composition in reverse to break our model into several subsystems. The decomposition structure is illustrated in Figure 5.2. In this figure, each of the boxes represents a separate machine. The three machines are `TRACK`, representing the trackside behaviour, `TRAINS` representing the train behaviour, and `COMMS` representing the communications layer used to send signals between the trackside and the trains. A line joining two boxes represents a synchronisation between machine operations. For example the `Check` operation of `TRAINS` is synchronised with the `DeliverMsg` operation of `COMMS`. An arrowhead indicates directed value-passing. The lines attached to just one box represent operations that are not involved in synchronisation.

Use of the communications layer models asynchronous signalling between the trackside and the trains. The trackside and the trains also synchronise directly through the `EnterSection` and `LeaveSection` operations. This models the fact that the trackside constrains train behaviour by determining precisely which section a train enters because the connectivity between sections is contained in `TRACK`. The synchronisation also models the fact that the trackside detects a train entering or leaving a section immediately when it happens. The guard of the composite `EnterSection` operation is the conjunction of the guards of the `EnterSection` operation of `TRACK` and the `EnterSection` operation of `TRAINS`, so that the composite operation is constrained by the operation in both machines. Occurrence of `EnterSection` in the `TRAINS` machine models a train actually entering a section, while occurrence of `EnterSection` in the `TRACK` machine models the detection of this by the trackside.

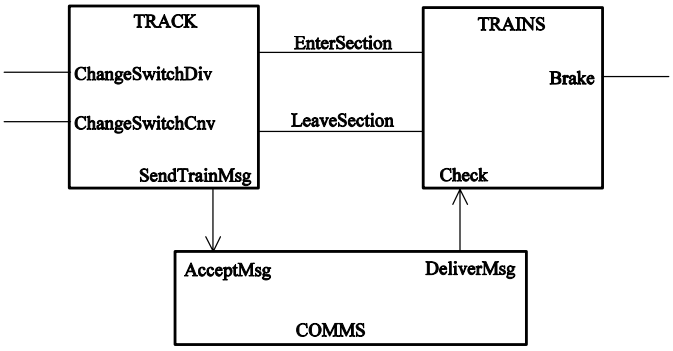


Figure 5.2: Decomposition structure.

The variables of the three machines are formed by partitioning the variables of the previous refinement. For example, the `tmsg` variable is placed in the `COMMS` machine, while the `braking` variable is placed in the `TRAINS` machine. A copy of the occupancy variables is placed in both `TRAINS` and `TRACKS`. These are not shared variables but instead they are separate variables which will always contain the same value since they are updated simultaneously with the same values by the shared `EnterSection` and `LeaveSection` operations. The values of the occupancy variables in the `TRAINS` machine models the actual positions of trains, while the values in the `TRACK` machines models the trackside's awareness of the positions of the trains.

```
REFINEMENT  Railway2  REFINES  Railway1
INCLUDES  track.TRACK, trains.TRAINS, comms.COMMS

OPERATIONS
…

Check(t) =
    PRE t ∈ TRAIN
    THEN
        VAR b ∈ BOOL WHERE
            b ← comms.DeliverMsg(t);
            trains.Check(t,b)
        END
    END;
```

Figure 5.3: Gluing the components together.

The decomposition of the railway is modelled in B by defining a B machine which glues together the three machines `TRACK`, `TRAINS` and `COMMS`. This gluing machine is declared to be a refinement of the previous model. This is partly shown in Figure 5.3 where the B `INCLUDES` clause is used to include each of the three machines in refinement `Railway3`. Figure 5.3 shows how the composite `Check` operation is formed by sequentially composing a call to the `DeliverMsg` operation of `COMMS` with the `Check` operation of `TRAINS`. The call to `DeliverMsg` returns a boolean value which is assigned to local variable `b`, and `b` is then passed on as an input parameter for the call to `Check` in `TRAINS`. The semantics of the B notation mean that the guard of the composite `Check` operation is the conjunction of the guard of `comms.AcceptMsg` and `trains.Check`, so, as well as modelling value passing, this construction also models synchronisation.

5.1.2 Event B Refinement

In event B each event of the abstract model is refined into a corresponding event of the concrete one. Informally speaking, a concrete event is said to refine its abstraction when

1) the guard of the former is stronger than that of the latter (guard strengthening), and

2) the gluing invariant is preserved by the conjoined action of both events.

Another frequent way of refining an event system consists in adding new events. This corresponds to observing the system with a finer granularity than in the abstraction. Such new events have an implicit (hidden) counterpart in the abstraction, namely, the event that does nothing. The new events that are introduced at some level must obey a specific constraint: they must not monopoly the "control" forever. In other words, they should not have the possibility to be fired indefinitely without letting the old events be executed from time to time. In practical terms, this means that, should the control be given exclusively (guards permitting) to the new events, then, at some point, the disjunction of their guards should become false.

A global constraint of a refined model with regards to its abstraction deals with deadlock. Since a system should normally run forever, then so must its refinement. But more generally, we require that a refined model should not deadlock more frequently than its corresponding abstraction.

Proofs of Correct Refinement

Suppose we have an abstract model with state v and invariant `I(v)`, and also a corresponding concrete model with state `w` and gluing invariant `J(v, w)`. If an abstract and corresponding concrete events are as follows:

```
ANY x WHERE
    P(x, v)
THEN
    v := E(x, v)
END
```

```
ANY y WHERE
    Q(y, w)
THEN
    w := F(y, w)
END
```

then the refinement is correct if:

$$I(v) \wedge J(v,w) \wedge Q(y,w) \Rightarrow \exists x.\ (\ P(x,v) \wedge J(E(x,v), F(y,w))\ )$$

This states that for each possible choice of the local variables of the concrete event there is a choice of the local variables of the abstraction that makes the gluing invariant (as modified by both events) true: the concrete event refines its abstraction. As can be seen the concrete guard is stronger than its abstract counterpart. In case of a new event of the form:

```
ANY y WHERE
    Q(y, w)
THEN
    w := F(y, w)
END
```

the correctness of refinement statement is simpler since that new event is only supposed to refine the non-event that does nothing. Formally:

```
I(v) ∧ J(v, w) ∧ Q(y, w)  ⇒  J(E(v, F(y, w))
```

Proofs of the Impossibility of Monopoly of New Events

Given a new event of the form

```
ANY y WHERE
    Q(y, w)
THEN
    w := F(y, w)
END
```

we must prove that, under the invariant, the event decreases a certain (natural number) variant expression `V(w)` that has to be exhibited. Notice that this variant expression must be the same for all the new events. This decreasing is thus a global property of the new events because we want to prevent their entire population from taking control forever. Should the variant be distinct for, say, two events, then they could very well stop individually after some steps, but the other could then take control in an endless ping-pong: this is clearly something we want to avoid. Formally:

```
I(v) ∧ J(v, w) ∧ Q(y, w)  ⇒  V(F(y, w)) < V(w)
```

Proofs of the Limitation of Deadlocks

For each abstract event of the form:

```
ANY y WHERE
    P(y, w)
THEN
    w := E(y, w)
END
```

and for an abstract invariant `I(v)` and a gluing invariant `J(v, w)`, the following statement must be proved:

```
I(v) ∧ J(v, w) ∧ P(x, v) ⇒ <disjunction of the concrete guards>
```

where the right hand side of the implication denotes the disjunction of the concrete guards. In other words, whenever an abstract event can be fired, then it must also be the case for a concrete event (not necessary the corresponding concrete event, it could be one of the new events). Notice that the guarding predicate `P(x, v)` that is on the left-hand side of the implication now corresponds to an abstract event. This contrasts with all the previous proof statements.

5.1.3 B-Action System Refinement

With the Event B refinement we can develop distributed systems. However, we cannot model more advanced features, like procedure refinement, with Event B. By using procedures in the distributed systems, we get a more structured system, as well as a general communication mechanism. B-action systems refinement supports also these features. Hence, the proof obligations generated for Event based B form a subset of the obligations generated for B-action systems. A more detailed description of proving the correctness of B-action systems can be found later in this chapter.

5.2 Iterative Construction of Gluing Invariants for Refinement

In this section we provide some guidance on constructing gluing invariants for proving an Event B-style refinement. We do this through an example refinement involving some timing constraints. The example is a simplification of a system based on the railway case study. We have an abstract model which specifies that within XX time units of being signalled to stop by the trackside system, a train must apply its brakes. To model the timing constraint, the abstract machine has a variable (clk) which models the current (discrete) time. It also has variables modelling whether the train has been signalled, the time at which it was signalled and whether the train is currently braking:

```
MACHINE stiming1

CONCRETE_CONSTANTS XX

PROPERTIES   XX ∈ NATURAL  ∧  XX > 0

ABSTRACT_VARIABLES clk , braking , signalled , stime

INVARIANT
     clk : NATURAL               /* current time                          */
     ∧
     stime ∈ ( 0..clk )          /* time at which a train is signalled    */
     ∧
     signalled ∈ BOOL            /* TRUE if the train has been signalled */
     ∧
     braking ∈ BOOL              /* TRUE if the train is braking          */
```

The timing constraint is modelled by adding the following constraint to the invariant:

```
( signalled=TRUE ∧ stime + XX ≤ clk ⇒ braking=TRUE )
```

This says that if the train has been signalled and the current time has reached the deadline stime+XX, then the train should be braking.

We introduce 3 operations in the machine to model the signalling of the train, the application of the brakes and the passage of time:

```
OPERATIONS

Signal =
     BEGIN signalled := TRUE   ||  stime := clk END;

Brake =
     BEGIN  braking := TRUE  END;
```

```
Tick =
     SELECT
          signalled=TRUE ∧ stime + XX ≤ clk + 1 ⇒ braking=TRUE
     THEN
          clk := clk + 1
     END

END
```

The timing constraint is imposed by preventing time from progressing unless the braking property would continue to be satisfied in the next time interval. In reality, we cannot prevent time from progressing, but we can satisfy the specification by ensuring that the timing constraint is met on time, *e.g.*, by applying the brakes of a train using the `Brake` event.

5.2.1 Refined Machine

Now we present a refinement of this in which the train is signalled by the use of an explicit message. We introduce 2 extra operations in the refinement corresponding to sending and receipt of a signal message. We assume that a message might get delayed or may get lost. To deal with this we also introduce a timeout mechanism.

We assume that the trackside system continually sends messages to a train every `MIT` time units. In this way a train knows by what time it should expect to receive a message. If it hasn't received a message within the expected time, it applies the emergency brakes. We also assume that there can be a delay in receiving messages and that the maximum expected delay is `MDT` time units. Thus a train should wait no longer than `MIT+MDT` time units in between message receipts before applying the brakes. The refinement has the following constants:

```
CONSTANTS  MDT, MIT, TIMEOUT

PROPERTIES

    MDT ∈ NATURAL ∧
    MIT ∈ NATURAL ∧
    TIMEOUT ∈ NATURAL ∧
    TIMEOUT > 0 ∧
    TIMEOUT ≤ MIT + MDT
```

The refinement has all the variables of the abstract machine as well as some new ones. Messages are delivered to the train via set `msgM` which is either empty or a singleton boolean value. Empty means there is no message pending for the train. A message of value `TRUE` means that the train has been signalled and should brake. When a message is received, its value is stored in `msgT`. We also introduce variables modelling the time at which the most recent message was sent and the time at which it was received.

```
ABSTRACT_VARIABLES
     clk , braking , signalled ,
     sgM , msgT , mstime , mrtime

INVARIANT
     msgM ∈ POW(BOOL)  ∧  card(msgM) ≤ 1  ∧  msgT ∈ BOOL  ∧
     mstime ∈ ( 0 .. clk )  ∧  mrtime ∈ ( 0 .. clk )
```

As well as the signal operation from the abstract machine, the concrete machine also has operations modelling the sending and receiving of messages:

```
SendMsg =
     SELECT
          msgM = ∅
     THEN
          msgM := {signalled}  ||
          msgS := signalled  ||
          mstime := clk
     END;

RecvMsg =
     ANY mm WHERE mm ∈ BOOL  ∧
          msgM = {mm}  ∧
          ¬(msgT = TRUE)
     THEN
          msgT := mm ||
          mrtime := clk ||
          msgM := ∅
     END;
```

The timing constraint on the concrete model says that if more than `TIMEOUT` time units have passed since the last message was received, of if `msgT` has the value `TRUE`, then the brakes should be applied. The timing constraint also says that if there is a pending message for the train and the train is not braking, then that message should be fresh, *i.e.,* no more than `MDT` time units should have passed since it was sent. The concrete `Tick` operation is thus defined as follows:

```
Tick =
     SELECT
          ( mrtime + TIMEOUT ≤ clk + 1 ⇒ braking=TRUE ) ∧
          ( msgT = TRUE ⇒ braking=TRUE ) ∧
          ( msgM/=∅ ∧ ¬(braking=TRUE) ⇒ clk + 1 ≤ mstime + MDT )
     THEN
          clk := clk + 1
     END
```

5.2.2 Gluing Invariant

The initial gluing invariant we use is simply the typing invariant give above. When we attempt to prove the refinement using the auto-prover and interactive prover of Atelier-B, we find that the following proof obligation (PO) is impossible to prove:

```
PRI > go(Tick.1)
    "`Local hypotheses'" &
    mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE &
    msgT$1 = TRUE => braking$1 = TRUE &
    not(msgM$1 = {}) & not(braking$1 = TRUE) =>
    clk$1+1<=mstime$1+MDT & stime+XX<=clk$1+1 &
    signalled$1 = TRUE &
    "`Check operation refinement - ref 4.4, 5.5'" &
    =>
    braking$1 = TRUE
```

This PO arises from the need to show refinement between the abstract and concrete `Tick` operations. We see from one of the antecedents that the goal (`braking=TRUE`) could easily be

shown if `msgT=TRUE`. So let us introduce a case split, on `msgT=TRUE`. In the interactive prover, we start a deduction and call the case split command:

```
PRI > dd
Starting Deduction
   braking$1 = TRUE

PRI > dc(msgT$1 = TRUE)
Starting Do Cases
   msgT$1 = TRUE => braking$1 = TRUE
```

Since this corresponds to a hypothesis, it is easily discharged using the autoprover, leading to the negative case:

```
PRI > pr
Starting Prover Call
   not(msgT$1 = TRUE) => braking$1 = TRUE
```

We add this antecedent to the hypothesis:

```
PRI > dd
Starting Deduction
   braking$1 = TRUE
```

It is also useful to add the negation of the goal to the hypothesis so that it can be used as an assumption in any clause that we add to the invariant. To do this, we perform a case split. The first case is trivially discharged, while the second provides a means to add the negated goal to the hypothesis:

```
PRI > dc(braking$1 = TRUE)
Starting Do Cases
   braking$1 = TRUE => braking$1 = TRUE

PRI > pr
Starting Prover Call
   not(braking$1 = TRUE) => braking$1 = TRUE

PRI > dd
Starting Deduction
   braking$1 = TRUE
```

The hypothesis list now includes the following:

```
mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE
stime+XX<=clk$1+1
signalled$1 = TRUE
not(braking$1 = TRUE)
not(msgT$1 = TRUE)
```

The goal could be proved by adding the following clause to the invariant

```
(  signalled=TRUE &
   not(braking=TRUE) &
   not(msgT = TRUE)
   =>
   mrtime + TIMEOUT <= clk + 1
)
```

A general rule of the thumb is that we should make use of as many of the existing hypotheses as possible before strengthening the invariant. This means that later POs will be less difficult to prove. In this case we can make use of the transitivity of inequality and make use of the second hypothesis above to weaken the clause to the following:

```
(   signalled=TRUE &
    not(braking=TRUE) &
    not(msgT = TRUE)
    =>
    mrtime + TIMEOUT <= stime + XX
)
```

We now save the current proof steps, as we need to exit the prover environment in order to add the clause to the invariant of the concrete machine. After adding the clause, we return to the interactive prover and reply the previous proof to the point at which we left it. We discharge the the additional clause (which now appears as a hypothesis) by applying modus ponens to it. This gives us

```
mrtime + TIMEOUT <= stime + XX
```

as a new hypothesis:

```
PRI > mh(signalled$1 = TRUE & not(braking$1 = TRUE) & not(msgT$1 =
TRUE) =>
                    mrtime$1+TIMEOUT<=stime+XX  )
Starting Modus Ponens on Hypothesis
        mrtime$1+TIMEOUT<=stime+XX &
        =>
        braking$1 = TRUE

PRI > dd
Starting Deduction
        braking$1 = TRUE
```

Our aim now is to prove this goal by discharging the hypothesis:

```
mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE
```

Thus we attempt to show that `mrtime$1+TIMEOUT<=clk$1+1` may be deduced from the current hypothesis. When we try to add this as a hypothesis, it becomes a sub-goal to be proven:

```
PRI > ah(mrtime$1+TIMEOUT<=clk$1+1)
Starting Add Hypothesis
        mrtime$1+TIMEOUT<=clk$1+1
```

To prove this sub-goal, we add 2 appropriate hypotheses to it and call the predicate prover:

```
PRI > ah(mrtime$1+TIMEOUT<=stime+XX)
Starting Add Hypothesis
        mrtime$1+TIMEOUT<=stime+XX => mrtime$1+TIMEOUT<=clk$1+1

PRI > ah(stime+XX<=clk$1+1)
Starting Add Hypothesis
        stime+XX<=clk$1+1 => (mrtime$1+TIMEOUT<=stime+XX =>
                mrtime$1+TIMEOUT<=clk$1+1)

PRI > pp(rp.0)
Starting Predicate Prover Call
```

```
Proved by the Predicate Prover
```

Having proved the sub-goal, we are left with the following main goal:

```
mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE
```

Since this is in the list of hypotheses, we simply call the auto-prover and we are done:

```
PRI > pr
Starting Prover Call
        "`Local hypotheses'" &
        mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE &
        msgT$1 = TRUE => braking$1 = TRUE &
        not(msgM$1    =    {})    &    not(braking$1    =    TRUE)    =>
clk$1+1<=mstime$1+MDT &
        stime+XX<=clk$1+1 &
        signalled$1 = TRUE &
        "`Check operation refinement - ref 4.4, 5.5'" &
        =>
        braking$1 = TRUE
```

The re-display of the original PO indicates that we have successfully proved it. Recall that to prove the PO, we needed to add the following clause to the invariant:

```
/* I1 - required to show refinement between both Tick ops: */

    (   signalled=TRUE &
        not(braking=TRUE) &
        not(msgT = TRUE)
        =>
        mrtime + TIMEOUT <= stime + XX
  )
```

5.2.3 Further Proof Obligations

Strengthening the invariant with this clause has the effect of introducing some new POs which themselves need to be proven. In this case we are left with one remaining PO that cannot be proven automatically or interactively:

```
PRI > go(RecvMsg.4)
        "`Local hypotheses'" &
        mm: BOOL &
        not(msgT$1 = TRUE) &
        msgM$1 = {mm} &
        not(braking$1 = TRUE) &
        not(mm = TRUE) &
        signalled$1 = TRUE &
        =>
        clk$1+TIMEOUT<=stime+XX
```

In this case, there are no hypotheses that can be used in an obvious way, we add this PO as a clause to the invariant. We make one simplification, which is to eliminate the variable `mm` and replace it by the value `FALSE` since the antecedent includes `not(mm = TRUE)`. Thus we further strengthen the invariant of the concrete machine with the following clause:

```
/*  I2 Required by RecvMsg to preserve I1 */

    (   signalled=TRUE &
        not(braking=TRUE) &
```

```
        not(msgT = TRUE) &
        msgM={FALSE}
        =>
        clk + TIMEOUT <= stime + XX
    )
```

Now when proving the PO *RecvMsg.4*, we need to add `msgM={FALSE}` as a hypothesis (the others are already given in the antecedent of the PO). This is easily achieved by making use of other hypotheses:

```
PRI > dd
Starting Deduction
        clk$1+TIMEOUT<=stime+XX

PRI > ah(msgM$1 = {FALSE})
Starting Add Hypothesis
        msgM$1 = {FALSE}

PRI > ah( msgM$1 = {mm} )
Starting Add Hypothesis
        msgM$1 = {mm} => msgM$1 = {FALSE}

PRI > ah( not(mm = TRUE) )
Starting Add Hypothesis
        not(mm = TRUE) => (msgM$1 = {mm} => msgM$1 = {FALSE})

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        msgM$1 = {FALSE} => clk$1+TIMEOUT<=stime+XX

PRI > dd
Starting Deduction
        clk$1+TIMEOUT<=stime+XX
```

We are now in a position to prove this goal by discharging the clause that we just added to the invariant:

```
PRI > mh( signalled$1 = TRUE & not(braking$1 = TRUE) & not(msgT$1 =
TRUE) &
          msgM$1 = {FALSE}  =>  clk$1+TIMEOUT<=stime+XX )
Starting Modus Ponens on Hypothesis
        clk$1+TIMEOUT<=stime+XX &
        =>
        clk$1+TIMEOUT<=stime+XX

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        "`Local hypotheses'" &
        mm: BOOL &
        not(msgT$1 = TRUE) &
        msgM$1 = {mm} &
        not(braking$1 = TRUE) &
        not(mm = TRUE) &
        signalled$1 = TRUE &
        =>
        clk$1+TIMEOUT<=stime+XX
```

This new invariant clause leads to further POs and there is one of these that is impossible to prove given the current invariant:

```
PRI > go(Tick.4)
        "`Local hypotheses'" &
        mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE &
        msgT$1 = TRUE => braking$1 = TRUE &
        not(msgM$1    =    {})    &    not(braking$1    =    TRUE)    =>
clk$1+1<=mstime$1+MDT &
        not(braking$1 = TRUE) &
        not(msgT$1 = TRUE) &
        signalled$1 = TRUE &
        msgM$1 = {FALSE} &
        =>
        clk$1+1+TIMEOUT<=stime+XX
```

The third clause of the antecedent for this PO can be discharged so that `clk$1+1<=mstime$1+MDT` can be used as a hypothesis:

```
PRI > dd
Starting Deduction
        clk$1+1+TIMEOUT<=stime+XX

PRI > ah(not(msgM$1 = {}))
Starting Add Hypothesis
        not(msgM$1 = {})

PRI > ah(msgM$1 = {FALSE})
Starting Add Hypothesis
        msgM$1 = {FALSE} => not(msgM$1 = {})

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        not(msgM$1 = {}) => clk$1+1+TIMEOUT<=stime+XX

PRI > dd
Starting Deduction
        clk$1+1+TIMEOUT<=stime+XX

PRI > mh(not(msgM$1 = {}) & not(braking$1 = TRUE) =>
clk$1+1<=mstime$1+MDT)
Starting Modus Ponens on Hypothesis
        clk$1+1<=mstime$1+MDT  =>  clk$1+1+TIMEOUT<=stime+XX
```

Now the antecedent of the final goal above allows us to deduce easily the following hypothesis:

```
clk$1+1+TIMEOUT <= mstime$1+MDT+TIMEOUT
```

This means that if we had as a hypothesis:

```
clk$1+MDT+TIMEOUT <= stime+XX
```

then the goal would be easily proven. Thus we construct the following clause to add to the invariant, with the above as a consequent, and many of the current hypotheses as antecedents:

```
/* I3 Required by Tick to preserve I2 */

    (   signalled=TRUE &
        not(braking=TRUE) &
        not(msgT = TRUE) &
        msgM={FALSE}
        =>
        mstime+MDT+TIMEOUT  <= stime + XX
    )
```

With this as a hypothesis, the proof of the PO can be completed:

```
PRI > mh(signalled$1 = TRUE & not(braking$1 = TRUE) & not(msgT$1 =
TRUE) &
         msgM$1 = {FALSE} =>  & mstime$1+MDT+TIMEOUT<=stime+XX )
Starting Modus Ponens on Hypothesis
        0<=XX-TIMEOUT-MDT+stime-mstime$1 &
        mstime$1+MDT+TIMEOUT<=stime+XX &
        =>
        clk$1+1<=mstime$1+MDT => clk$1+1+TIMEOUT<=stime+XX

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        "`Local hypotheses'" &
        mrtime$1+TIMEOUT<=clk$1+1 => braking$1 = TRUE &
        msgT$1 = TRUE => braking$1 = TRUE &
        not(msgM$1 = {}) & not(braking$1 = TRUE) =>
clk$1+1<=mstime$1+MDT &
        not(braking$1 = TRUE) &
        not(msgT$1 = TRUE) &
        signalled$1 = TRUE &
        msgM$1 = {FALSE} &
        =>
        clk$1+1+TIMEOUT<=stime+XX
```

Fortunately, we now reach a stage where all the newly introduced POs can be proven with the current invariant and do not require further strengthening of the invariant.

5.2.4 Constant Properties

As well as strengthening the invariant, proving the refinement POs also requires that we introduce some constraints between the constant `XX` used in the abstract and the constants `TIMEOUT` and `MDT` used in the refinement. Such as constraint is required when proving refinement between the `Signal` operations:

```
Example of how property such as TIMEOUT+MDT+1<=XX required:

PRI > go(Signal.3)
        "`Local hypotheses'" &
        not(braking$1 = TRUE) &
        not(msgT$1 = TRUE) &
        msgM$1 = {FALSE} &
        =>
        mstime$1+MDT+TIMEOUT<=clk$1+XX

PRI > dd
Starting Deduction
        mstime$1+MDT+TIMEOUT<=clk$1+XX
```

Now, we have as a hypothesis that `mstime$1<=clk$1` (this is easily discovered by performing a search for hypotheses matching mstime$1) so we add this:

```
PRI > ah(mstime$1<=clk$1)
Starting Add Hypothesis
        mstime$1<=clk$1 => mstime$1+MDT+TIMEOUT<=clk$1+XX
```

Now, we see that we require that

```
TIMEOUT+MDT+1<=XX
```

so we add this constraint to the PROPERTIES clause of the concrete machine. Now we can complete the proof of the PO by adding this property to the goal and calling the predicate prover:

```
PRI > ah(TIMEOUT+MDT+1<=XX)
Starting Add Hypothesis
        TIMEOUT+MDT+1<=XX => (mstime$1<=clk$1 =>
mstime$1+MDT+TIMEOUT<=clk$1+XX)

PRI > pp(rp.0)
Starting Predicate Prover Call
Proved by the Predicate Prover
        "`Local hypotheses'" &
        not(braking$1 = TRUE) &
        not(msgT$1 = TRUE) &
        msgM$1 = {FALSE} &
        =>
        mstime$1+MDT+TIMEOUT<=clk$1+XX
```

The above examples illustrate how the prover can be used to construct additional clauses for the gluing invariant in order to prove refinement POs. Using the prover to guide the construction of the gluing invariant means we are less likely to guess invariants that are too strong or that contain errors.

5.3 Guidance on Proving Loops

5.3.1 Overview

Loops are unavoidable control structures in programming, and it is worthwhile to separate their development from the development of purely sequential operations. To do so, we issue the following recommendations:

- it is necessary to identify in implementation operations subparts which are implemented with loops, to isolate them, and to encapsulate them in dedicated operations (local operations or operations from imported abstract machines) that define them independently from their future implementation.

- the structure of the operation implementing a loop should be as simple as possible, loops should never be nested, and their body should be essentially an operation call, perhaps some conditional substitution. If efficiency considerations allow it, loops bodies should modify only local variables of the operation whose values are assigned to implementation variables or operation results after the loop. For instance, the implementation.

During the modelling phase, it is not mandatory to formulate completely the loop invariant, however, for type checking purposes, it should include at least a predicate, the typing predicate of the loop index for example. For the same reason a well typed variant should be present. The complete definition of the invariant and variant of the loop will be carried out at detailed design.

5.3.2 Loop Development

The starting point of the development of a loop is the operation which defines abstractly the final effects that it must achieve. If we suppose that this operation have been defined in an abstract machine, then, in the refinement preceding the implementation including the loop, it is recommended to reformulate the abstract definition of the loop with a before/after substitution which uses the variables of the implementation and states the properties of the individual elements that will be treated in the loop. Usually, these properties involve quantified variables ranging over the ordered sets that will be scanned in the loop, or domain and co-domain restrictions over such ordered sets, etc.

It is not always possible to reformulate in a simple manner an abstract definition. In those cases it is preferable to leave the abstract definition as it is.

5.3.3 Loop Invariants

The identification of loop invariants, is one of the most challenging activities of formal development. Below is an approach to obtain loop invariants by systematic transformations of abstract definitions of loops.

Starting from the predicate of the last abstract definition, the loop invariant is constructed as follows:

1. To indicate that the expected result (the abstract definition) is achieved progressively at each loop iteration, one or more constants of the predicate are replaced by loop indices. The constants to be replaced are selected according to the scanning order of the elements in the loop

2. As it is the implementation variables that are constructed progressively in the loop, the "after" variables of the predicate of the abstract definition are replaced by the implementation variables

3. The "before" variables of the predicate of the abstract definition correspond either to variables of the implemented module or to input parameters. Therefore, the "before" variables corresponding to abstract variables are decorated by the $0.

4. The predicates of the implementation invariant concerned with the abstract and concrete variables involved in the predicate produced by the previous steps are modified in such a manner that they indicate the range of scanned values over which they are still valid; the resulting predicates are added to the loop invariant.

5. It may be necessary to add to the loop invariant predicates assuring that the pre-conditions of the operations invoked within the loop body hold.

6. All the variables updated in the loop that appear in the predicate resulting from steps 1-5 are typed, and the typing definitions are added to the loop invariant.

Introducing an intermediate refinement, which reformulates the abstract definition in terms of indices covering intervals, is sometimes worthwhile, as this formulation is more adequate for loop invariants. Furthermore, this intermediate refinement splits the proof of the loop in two steps. The first step proves that the abstract definition is actually refined by the reformulated

abstract definition. The second step is mainly concerned with the specific loop proof obligations (loop initialization establishes the loop invariant, loop body preserves the loop invariant, and loop variant decreases), as the proof that the loop actually refines the reformulated abstract definition has been considerably simplified by the fact that the loop invariant has been derived from the reformulated definition.

5.4 Refining from System Level to Software Level Leading into Decomposition

System decomposition is achieved by *decomposing* the overall model into a number of sub-models. For instance, in a train system, we could have some control parts replicated and embarked within trains and other control parts replicated and installed along the track, and perhaps, yet another control part at the extremity of some line, etc.

The role of decomposition is clear. Once separated from the main body, a sub-model can be developed further independently from the rest of the system, which becomes its, so-called, *environment*. Notice that this notion of is essentially *relative*: each sub-model (or group of sub-models) being the environment of others and vice-versa.

There are two kinds of decompositions:

- The decomposition of a system into several subsystems: for instance, in the railway case study, the railway system is divided into two subsystems: the "on-board" subsystems (that makes the train brake) and the "wayside" subsystems (that makes the switch move).

- The decomposition of a (sub)system into its physical and controller part. For instance, the on-board subsystem is divided into its physical part (the B model of the train, with its speed, brakes, *etc.*) and its controller part (the B model of the software, that deals with measured speed, brake commands, *etc.*)

The question is, which kind of decomposition should come first. If the system is decomposed into its physical and software parts, and then each part is decomposed in two subsystems (like the diagram below), there might be incoherence. For example, a system may have three controllers for only two subsystems.

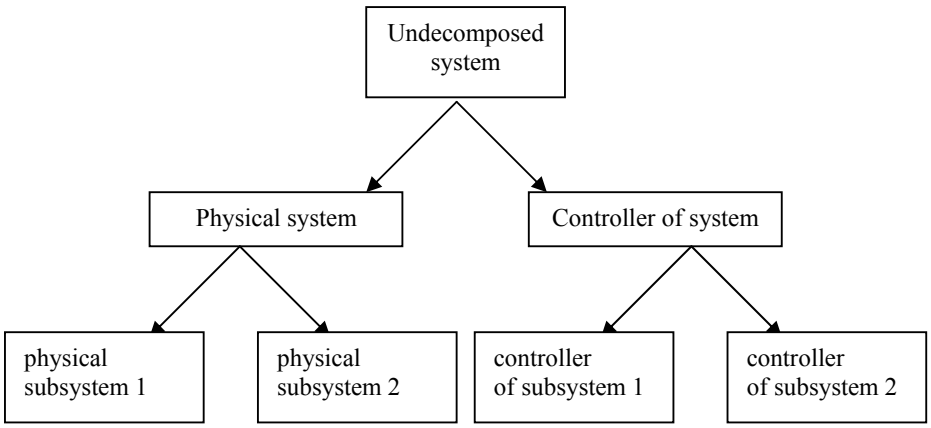


Figure 5.4: Bad example of decomposition.

On the contrary, with the following diagram, the system decomposition will not create such incoherencies.

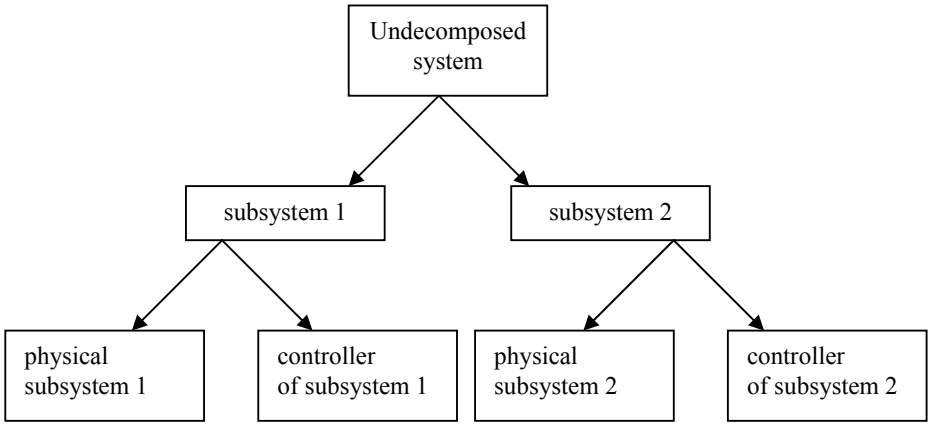


Figure 5.5: Good example of decomposition.

Decomposition in B

A system `C_comp` can be decomposed into two components `A_comp` and `B_comp`. Let us consider the abstract machine `C_comp` below, where `x` and `y` are local variables and `u`, `v`, and `z` are global variables. Furthermore, we assume that we have operation `A_oper` that refers to the variables `x` and `u` as well as `B_oper` that refers to `y` and `v`.

```
MACHINE  C_comp
INCLUDES
   Global_u, Global_v, Global_z
VARIABLES
   x, y
INVARIANT
   Inv_A ∧ Inv_B
INITIALISATION
  x := x0 || y := y0
OPERATIONS
   A_oper =  SELECT P  THEN S END;
   B_oper =  SELECT Q  THEN T END
END
```

The *parallel decomposition* of the system `C_comp` into the components `C'_comp`, `A_comp` and `B_comp` is then defined by splitting the variables, procedures and operations of `C_comp` as follows.

```
MACHINE  C'_comp
INCLUDES
  A_comp, B_comp
END
```

```
MACHINE  A_comp
INCLUDES
    Global_u, Global_z
VARIABLES
   x
INVARIANT
   Inv_A
INITIALISATION
  x := x0
OPERATIONS
   A_oper =  SELECT P  THEN S END
END
```

```
MACHINE  B_comp
INCLUDES
   Global_v
SEES
   Global_z
VARIABLES
   y
INVARIANT
  Inv_B
INITIALISATION
 y:= y0
OPERATIONS
   B_oper =  SELECT Q  THEN T END
END
```

After the decomposition `A_comp` contains the variable `x` and `u` while `B_comp` contains the variables `y` and `v`. The global variable `z` is a common variable of `A_comp` and `B_comp`. Due to B restrictions it can only be changed (included) by one of the components. The invariant, the initialisation, the procedures , as well as the operations referring to the variables `x` and `u` of `A_comp` are included in `A_comp`, while the ones referring to the variables of `B_comp` are included in `B_comp`. This rule can be applied in reverse. In that case it is called *parallel composition* [Butler96].

5.4.1 Healthcare Case Study

A control system, such as the Fillwell workstation in our case study, consists of four components: a *controller*, *sensors*, *actuators*, and a *plant*. These components are depicted in Figure 5.6.

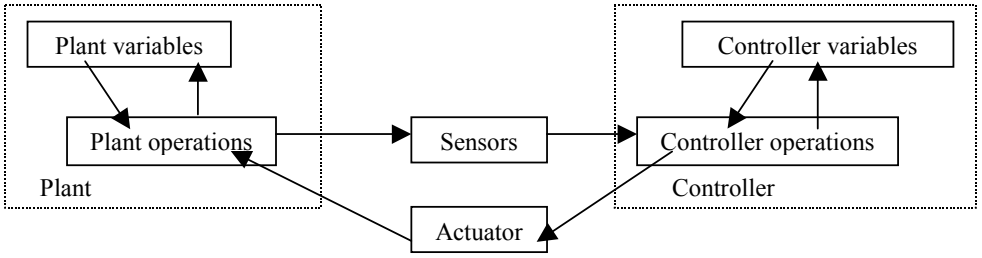


Figure 5.6: The structure of the control system specification

The purpose of the controller in a control system is to ensure that the plant operates within the predefined requirements. For simplicity, we consider here only controllers with *responsive*

dynamics. Such controllers become active only when they observe a discrepancy, or some other important event, in the plant.

As for the plant in a control system, we consider it here to have *autonomous dynamics*. This means that the plant changes its state by itself and, in a way, independently of the controller. Thus, even without the controller, the plant would keep on doing something by itself.

The purpose of the sensors and actuators then is to act as information *converters* between the controller and the plant. The sensors convert *measurements* from the plant into *readings* for the controller. Correspondingly, the actuators convert *commands* from the controller into *control signals* to the plant.

The difficulty in the development of a control system is in finding a suitable combination of control logic, sensors, and actuators that allow the controller to maintain the state of the plant within predefined tolerances. Clearly different choices affect the optimality of the overall performance. However, as a general guideline, one tries to minimise manufacturing costs and maximise reliability. This leads to the development of intelligent control logic to allow the use of simple and reliable sensors and actuators. From this point of view, the use of a co-design method, *i.e.*, a method that allows postponing of hardware details during the development process, supports these general development goals. As an example of postponing hardware details, we first consider a robot arm movement generally and only later introduce motor specific details, like acceleration and deceleration.

When thinking of control systems appearing as part of healthcare systems, another difficulty is that such systems should exhibit provable safety and reliability properties. Thus, during the development of a control system, one has to be able to incorporate safety and reliability measures into the system in such a way that these measures can be demonstrated to the customer when necessary. Typically this means increasing complexity in the control logic, as it has to cope with malfunctioning sensors and actuators, along with duplication of some physical sensors and, sometimes, even actuators.

5.4.1.1 Determining Controller and Plant

When enough implementation details are introduced, we proceed by decomposing the component specifications into controller and plant specification pairs. Thus, for each component we obtain both a controller specification and a plant specification. The obtained plant specification describes autonomous behaviour of the component, whereas the controller specification describes algorithms that guide the plant behaviour. Thus, the role of the controller is to react to changes in the plant. We can note that the parallel decomposition of a B-action system into several B-action systems that have the same effect as the initial B-action system is only a structuring technique.

In B-action systems terms, the plant machine contains the transitions to the new states, but before each such transition takes place, a controller machine procedure is called for handling the required decisions. If the controller machine needs not intervene, the procedure body is modelled with a stuttering operation `skip`. Thus, the procedures of the controller machine are thought of as interrupt procedures.

To obtain such a decomposition in B-action systems, the operations in the machines to be decomposed have to be split, and the variables are partitioned among the plant and the controller machines. As for the splitting of operations, each operation of the form

```
Oper  =  SELECT state=act1 ∧ A THEN B; state := act2 END
```

is replaced with the operation in the plant

```
Oper'  =  SELECT state=act1 ∧ A THEN Act2 END,
```

where the procedure `Act2` of the controller is

```
Act2  =  PRE state=act1 ∧ A  THEN B; state:=act2 END.
```

Obviously, the effect of the new operation `Oper'` is the same as the effect of the old operation `Oper`. Adding procedures and keeping the old functionality agrees with the superposition refinement step.

As an example, let us consider the operation for delivering a plate. When the operating table with a plate is at the lower end position and the Robot is ready to take the plate `ReadyFor_blank`, the plate is removed (`blank := FALSE`) and the liquid level is set to zero. Also the variable modelling the Analyser waiting for the Robot is reset. The operation `Deliver` in the B-action system to be decomposed,

```
Deliver = SELECT astate1=aidle1 ∧ acmd=deliver ∧ plate=TRUE ∧ blank = TRUE ∧
                 zpos=zmin
          THEN
                 ReadyFor_blank;
                 ChangeBlankAndResetPlateLiq(FALSE);awaited:=FALSE;
                 astate1:=adeliver1
          END,
```

is replaced with the operation `Deliver` in the plant,

```
Deliver = SELECT astate1=aidle1 ∧ acmd=deliver ∧ plate=TRUE ∧ blank = TRUE ∧
                 zpos=zmin
          THEN
                 ReadyFor_blank; Ddeliver
          END,
```

where the procedure `Ddeliver` of the controller is:

```
Ddeliver =    PRE astate1=aidle1 ∧ acmd=deliver ∧ blank=TRUE ∧ plate=TRUE ∧
                  zpos=zmin
              THEN
                  ChangeBlankAndResetPlateLiq(FALSE) || awaited:=FALSE ||
                  astate1:=adeliver1
              END.
```

We can note that the procedure call on the imported procedure `ReadyFor_blank` in the operation `Deliver` remains in the plant. It can be seen as a procedure call to the controller of the system that exports the procedure. Since the procedures are given in an abstract machine construct, we have multiple assignments instead of sequential composition of assignments in the procedure `Ddeliver`.

5.4.1.2 Determining Sensors and Actuators

In the last development step, we determine the sensors and the actuators for the components from the controller and plant specifications. In particular, the variables of the plant and the controller specification usually model more than just properties of these entities; some of these variables are used for communication. From these variables, we find sensors and actuators that are used to synchronise the plant specification with the controller specification.

In B-action systems, the actuators and sensors are global variables of the plant and the controller, and these variables are put into separate machines. The sensor variables are set by the plant and read by the controller, while the actuator variables are set by the controller and read by the plant. Due to this, the plant INCLUDES the sensor variables and SEES the actuator variables. Dually, the controller INCLUDES the actuator variables and SEES the sensor variables. An invariant relates the state of the whole control system to the plant and the controller state, as well as to the sensor variables and to the actuator variables.

The operations of the plant refine the corresponding operations of the initial specification. The general decomposition schema using B-action systems [Sekerinski98a] is shown in Figure 5.7**.** Encapsulating the actuator and sensor variables in separate machines allows abstracting from the details of a particular communication mechanism, which depends on the underlying hardware and operating environment. This allows late decisions regarding the hardware implementation.

```
REFINEMENT  MachinePlant
REFINES  Component_Ref
INCLUDES  MachineCtrl, MachineSensors
VARIABLES
 <plant state>
INVARIANT
 <variable types> ∧
 <refinement invariant for plant attributes>,
 <controller attributes, actuators, sensors>
INITIALISATION
 <plant state and sensors initialisation>
OPERATIONS
 Pl_oper = SELECT Pl_oper_guard
    THEN plant state and
         sensor change;
         Ctrl_oper(parameters)
    END;
 …
END;
```

```
MACHINE  MachineCtrl
INCLUDES  MachineActuators
SEES  MachineSensors
VARIABLES
 <controller state>
INVARIANT
 <variable types>
INITIALISATION
 <controller state and actuators
  initialisation>
OPERATIONS
 Ctrl_oper(parameters) =
    PRE  Pl_oper_guard
    THEN <controller state and
         actuator change>
    END;
 …
END
```

```
MACHINE MachineSensors
VARIABLES
 sensors
INVARIANT
 <sensors types>
INITIALISATION
 sensors :∈ <any value>
OPERATIONS
 SetSensor(ss)  =  sensor := ss;
 …
END
```

```
MACHINE MachineActuators
VARIABLES
 actuators
INVARIANT
 <actuators types>
INITIALISATION
 actuators :∈ <any value>
OPERATIONS
 SetActuator(aa)= actuator := aa;
 …
END
```

Figure 5.7: Decomposition schema for control systems within B-action system formalism.

5.4.2 Transportation Case Study

Track decomposition

The line is divided into sectors (from 1 to *n*). Each sector is connected to one or two sector(s), called its adjacent sector(s). A switch is a special sector that has either two incoming or two outgoing sectors. The connectedness relation for the sectors is reflexive.

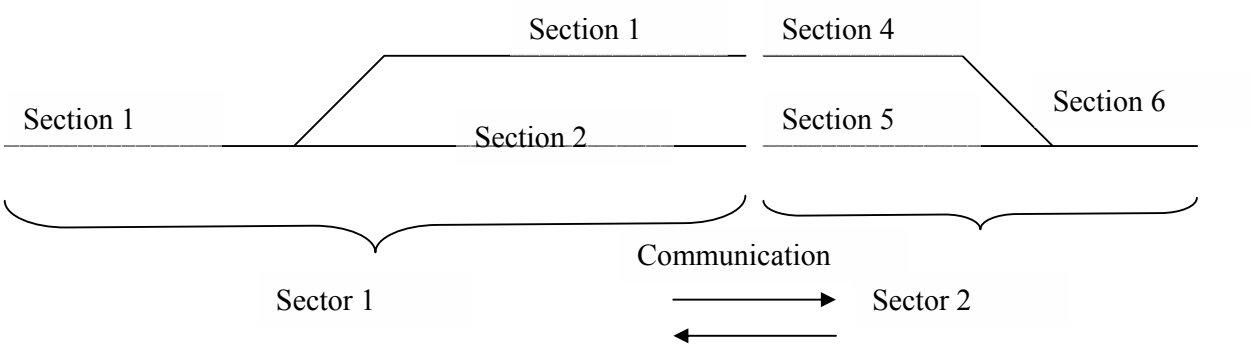
Each sector is associated with one piece of software, which controls this sector (position of switches, emission of messages to train and to adjacent sector, and so on).

Singularities, such as switches, stopping points, and sections, are located on the line, and therefore are presented on sectors. These singularities are represented in the software by arrays of constant data. In fact, this software is divided into two parts:

- A *generic* part which is the same for all sectors, deals with the algorithmic nature of the software.
- A specific part that is sector-dependent.

Each sector is divided into sections.

Adjacent sectors communicate: each sector sends and receives messages from its adjacent sector(s). These messages are essentially authorization to enter a section.



A (simplified) message between sectors is composed of:

- The number of the section that emitted the message;
- A boolean (authorization or not, for a train in the adjacent sector, to enter the section that emitted the message)

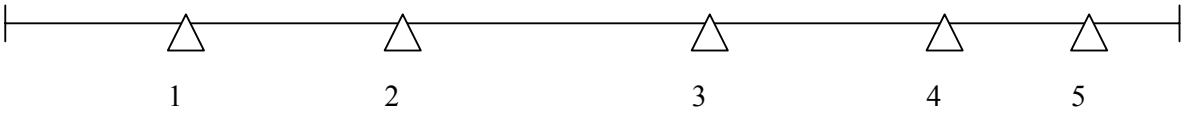
Sections also transmit messages to trains. Each section emits different messages, but if there are several trains on a single section, they will receive the same message. A (simplified) message to a train is composed of:

- The number of the section that emitted the message;
- A boolean (authorization or not to go forward)

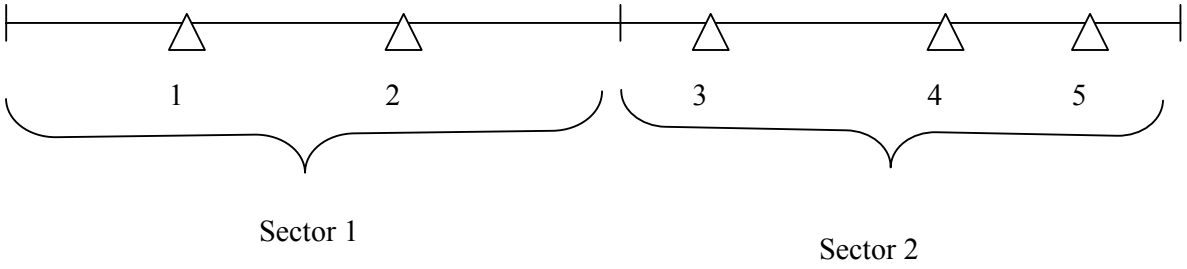
If a train does not receive any message over a 5 second period (14 cycles), then the onboard equipment considers that there is no authorisation to enter an adjacent sector. Similarly, if a sector does not receive any message during 5 second period, then the sector equipment considers that there is no authorisation to enter an adjacent sector.

Description of the invariants

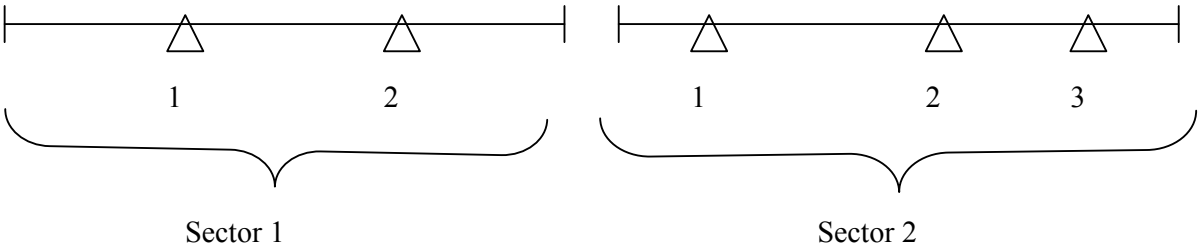
Here is a simplified example with absolute and relative identification of stopping points. Below is the line invariant, with absolute identification number of each stopping point:



Now, the line is decomposed into two sectors:



Now, we obtain the invariants of the two sectors, with relative identification numbers for each stopping point:



with the gluing bijection defined according to Table 4.

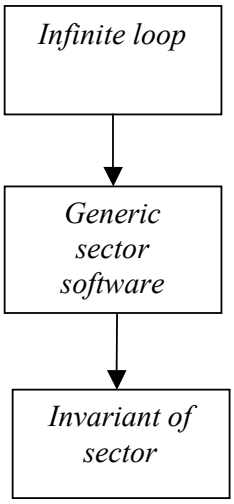
Absolute Stopping Point	Sector	Relative Stopping Point
1	1	1
2	1	2
3	2	1
4	2	2
5	2	3

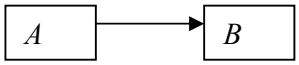
Table 4: Gluing bijection.

The software is linked with the relative identification, whereas in the system model, both identifications are present, with the gluing relation.

Software architecture

The current software architecture is described in . For each sector, an infinite loop calls the generic software, which is linked with the invariant of the sector.

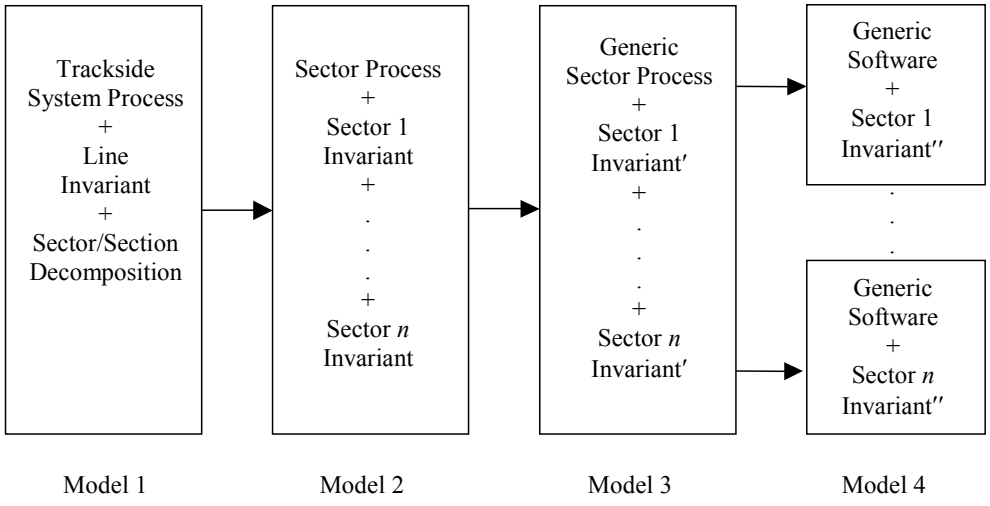




Means that A have access to B through the B clause *contain*.

The track system

In this schematic, the invariant term means "constant data describing the line/sector" and is not related to the invariant clause of a B component

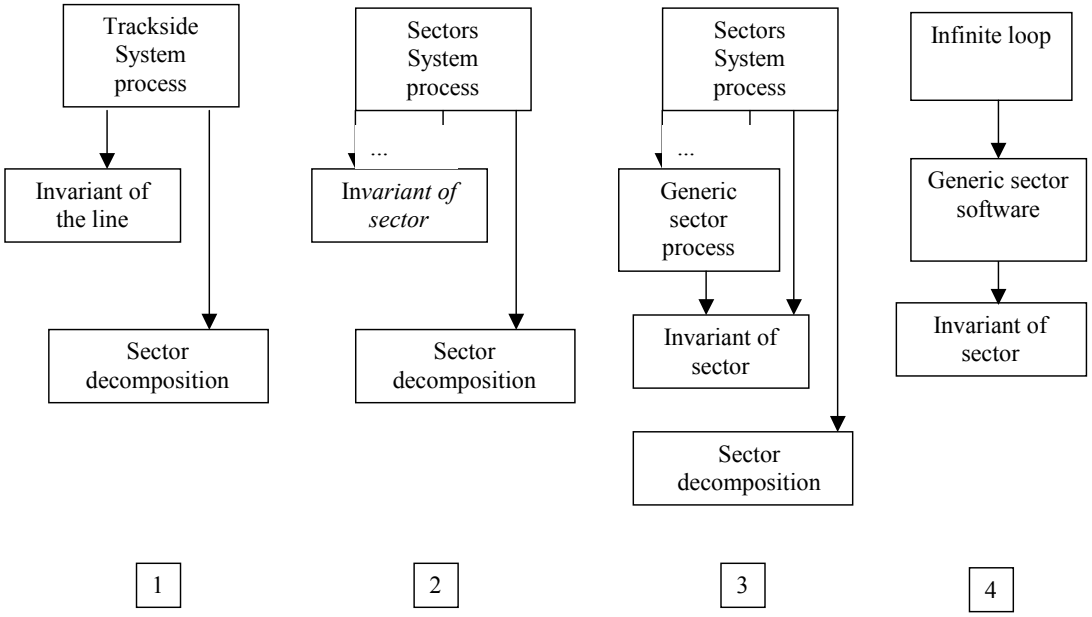


In model 1, the proof of the correct decomposition is performed (basically, $line = \bigcup_{1}^{n} sector$).

In model 2, sector process is introduced. This sector process includes communication (message from a sector to another). the sector process is a refinement of the system process.

In model 3, the sector process is refined as a generic sector process. Invariant' are linked with Invariant by a gluing relation.
In model 4, the generic sector process is refined into a generic software. Invariant'' are linked with Invariant' by a gluing relation.

Each model is related to the system development shown in next.



In model 2, "sectors system process" has access to the n "sector invariant".
In model 3, "sectors system process" has access to the n "sector invariant", and to the n "generic sector process". Each "generic sector process" has access to its related invariant.
In model 4, for each sector, an infinite loop calls the generic software, which is linked with the invariant of the sector.

5.5 Safety Requirements in Refinement

In the healthcare case study, after translating the basic statechart diagram (Figure 3.6) into a B-action system, we gradually refine the system towards a B-model. For the refinement process we identify the attributes suggested in the use case descriptions. These attributes/variables are then added gradually to the specification making it more concrete. At each refinement step we add one or more variables. For each variable we state the safety conditions and the properties of the variable that can be found in the requirements specification. We add the computation concerning the new variable(s) to the existing operations. New operations that only assign the new variable(s) may also be added. A particular refinement method consists in adding new functionality to a specification in this way preserving the old behaviour. This type of refinement

is referred to as superposition refinement. When dealing with complex control systems it is especially convenient to stepwise introduce details about the system to the specification and not to have to handle all the implementation issues at once.

The refinement `Component_Ref` (Figure 5.8) of the machine `Component` (see Figure 3.8 in Section 3.1.2.5) can be derived from the state diagram in Figure 5.10. Some of the variables in `variable_list` in the `VARIABLES` clause are additional to those in the abstract machine `Component`. The types of the variables, their properties and internal relations as well as their safety requirements are given in the `INVARIANT` clause. The variables are updated in the operations. Compared to the abstract machine specification `Component` the refinement machine has more detailed operations. The guards have more precise conditions under which the operation can be performed and also the newly introduced variables can be changed in the operation. In the operation `Service1` the new variable(s) are referenced to in the predicate `Service1_prec` as well as in the substitutions in `Service1_comp`.

Furthermore, new operations concerning each service/subroutine may be introduced that take care of new features introduced by the new variables. The exported procedures are incorporated in the system as before.

5.5.1 Safety Issues During the Refinement

As the system development proceeds we obtain more elaborated information about faults and conditions of failure occurrence. The refinement step introduces distinction between faults. The operation `Service1_fail` models a fault resulting from an attempt to provide a service from an incorrect initial state:

```
Service1_fail = SELECT ¬Service1_prec  ∧ state = Idle THEN state := Suspension1 END
```

This situation might be caused by faults occurred previously or by a logical error in the calling command. The operation `Service1_notok` models fault occurrence during the execution of the action:

```
Service1_notok = SELECT ¬Service1_postc ∧ state=Service1 THEN state := Suspension1 END
```

These kinds of faults are caused by the physical failures of the system components involved in the execution.

```
REFINEMENT  Component_Ref
REFINES  Component
VARIABLES  variable_list
INVARIANT
   variable_types  ∧
   variable_relations ∧
   variable_requirements
INITIALISATION  variable_initialisation
OPERATIONS
   New_command(ss) =
      PRE ss ∈ {serv1,serv2,serv3}
      THEN
         SELECT  state = Idle THEN cmd := ss END
      END;
   Service1 =
```

```
        SELECT  Service1_prec ∧ state = Idle
        THEN
            state := Service1 || Service1_comp
        END;
    Service1_fail =
        SELECT  ¬Service1_prec ∧ state = Idle
        THEN
            state := Suspension1
        END;
    Service1_ok =
        SELECT  Service1_postc ∧ state = Service1
        THEN
            state := Idle || Service1_endcomp
        END;
    Service1_notok =
        SELECT  ¬Service1_postc ∧ state = Service1
        THEN
            state := Suspension1
        END;
    Remedy1 =
        SELECT  state = Suspension1
        THEN
            fix _the_error1  || state := Service1
        END;
    …
    Services_notready =
        SELECT  state = Idle
        THEN
            state := Suspension
        END;
    Remedy =
        SELECT state = Suspension
        THEN
            fix _the_error  || state := Idle
        END;
    Failure =
        SELECT state ∈ {Suspension, Suspension1, Suspension2, Suspension3}
        THEN
            state := Abort
        END
END
```

Figure 5.8: A refinement of a B-action system.

We also introduce a distinction between different repair procedures by adjusting the `Remedy` operation for each fault accordingly.

The action `Failure` models system shut down:

```
Failure = SELECT state ∈ {Suspension,Suspension1,…,Suspension3}
          THEN state := Abort END
```

The action becomes enabled when the attempts to fix occurred faults fail.

Safety analysis proceeds by an identification of different failure modes of the system. It is conducted by considering the consequences of multiple faults. In the system specification we proceed by specifying the status of each component (failed or functioning), specifying system failure modes and finally by introducing error messages.

5.5.2 Proving the Correctness of the Refinement

Within Atelier-B we can actually prove formally that the refinement is sound. For this a number of proof obligations first need to be generated. Let us assume that we have two systems `A_comp` and `C_comp` as below. The variable `x` is a variable of both `A_comp` and `C_comp`. In the refinement step the variable `y` is a new variable of `C_comp`. The operations `A1`, `A2` and `A3` are changed to also take `y` into account in `C_comp`. The operations `B1` and `B2` are created during the refinement step.

```
MACHINE  A_comp
VARIABLES
 x
INVARIANT
 Inv_A
INITIALISATION
 x := x0
OPERATIONS
 A1 = SELECT P1 THEN S1 END;

 A2 = SELECT P2 THEN S2 END;

 A3 =  SELECT P3 THEN S3 END
END
```

```
REFINEMENT  C_comp
REFINES A_comp
VARIABLES
 x, y
INVARIANT
 Inv_C
INITIALISATION
 x := x0 || y:= y0
OPERATIONS
 A1 = SELECT P1' THEN S1' END;

 A2 = SELECT P2' THEN S2' END;

 A3 = SELECT P3' THEN S3' END;

 B1 = SELECT Q1 THEN T1 END;

 B2 = SELECT Q2 THEN T2 END
END
```

The proof obligations needed for proving that a concrete system `C_comp` is a correct refinement of a more abstract system `A_comp` are as follows [Back96, Walden98a]. For each proof obligation the corresponding proof obligation in Event B is given in the parentheses.

1. The initialization in `C_comp` should be a refinement of the initialization in `A_comp`, and the initialization should establish the invariant `Inv_C`. (Proofs of Correct Refinement)

2. Each operation `Ai` (`i=1,2,3`) in `C_comp` should refine the corresponding operation (the operations with the same name) in `A_comp`, and they should preserve the invariant `Inv_C`. (Proofs of Correct Refinement)

3. Each new operation `Bj` (`j=1,2`) `in C_comp` (that do not have a corresponding operation in `A_comp`) should only concern the new variable `y`, and preserve the invariant `Inv_C`. (Proofs of Correct Refinement)

4. The new operations `Bj` (`j=1,2`) in `C_comp` should terminate, if they are executed in isolation. (Proofs of the Impossibility of Monopoly of New Events)

5. Whenever an operation `Ai` (`i=1,2,3`) in `A_comp` is enabled, either the corresponding operation `Ai` in `C_comp` is enabled or then one of the new operations `Bj` (j=1,2) in `C_comp` is enabled. (Proofs of the Limitation of Deadlocks)

6. Whenever a fault occurs in A_comp (a fault-operation in A_comp is enabled), a fault could also occur in C_comp (a fault-operation in C_comp could be enabled). (Proofs of the Limitation of Deadlocks)

With the fault-operations in (6) we mean the operations leading to a *suspension* state. Hence, according to (6) a general fault is partitioned into distinct faults during the refinement process.

The proof obligations (1)–(3) above are automatically generated by Atelier-B. Proof obligations (5) and (6) can be generated automatically if an *exit*- as well as an *exitFail*-operation is added to the system. The *exit*-operation should be enabled when all the other non-failure operations in the system are not enabled. In A_comp the *exit*-operation would be

SELECT ¬P1 ∧ ¬P2 ∧ ¬P3 **THEN** skip **END**.

The *exitFail*-operation can be created similarly with the guards of the failure-operations. The proof obligation (4) requires a variant that is decreased by the new operations. Using Atelier-B we need to introduce some additional machine constructs discussed in [Walden98b,]. With the help of the Evt2b translator [ClearSy01] also proof obligation (4) can be generated automatically as discussed in [Butler96].

During the refinement process global procedures, procedures that are declared in one system and called from other systems as P above, are refined in the same way as the operations of the system. Hence, let us assume that we have an abstract system A_comp with a global procedure P and a refined system C_comp with the refined procedure P′. Then for C_comp to be a refinement of A_comp, the six proof obligations concerning refinement of distributed systems given earlier should hold, as well as:

7. Procedure P in A_comp should be refined by the corresponding procedure P′ in C_comp, and P′ should also preserve the invariant.

8. If procedure P is enabled so should P′ be, or then the operations in C_comp should enable P′.

Proof obligation (7) is automatically generated by Atelier-B and corresponds to the proof obligations concerned with Correct Refinement. Proof obligation (8) requires some extra constructs to be created. It has to be proven that when executing the operations they will eventually enable the refined global procedure P′ when the abstract procedure P is enabled. Hence, we should provide a variant that is decreased by each operation. For more details on procedures and their proof obligations, see [Walden98b].

The proof obligations can be discharged with the help of the auto-prover and interactive prover in Atelier-B. The auto-prover uses a database of proof rules in order to perform the proofs. In case the needed proof rules are not found in this database, some new rules can be added via the interactive prover of Atelier-B. The proof obligations that were not proved automatically can be proven interactively with the interactive prover.

5.6 Refinement with UML+B

In order to get a graphical interface to the formal method to facilitate the reading and analysing of the development, UML artefacts are integrated in the formal development process. The stepwise introduction of implementation details (features) can be applied by adding these features into class and statechart diagrams. Part of the refined diagrams is automatically generated from the more abstract ones. New variables and operations can be added to the class diagrams. The more concrete behaviour of the system can then be modelled with new states and more complex transitions in the statechart diagrams taking into consideration the new variables and operations of the corresponding class diagram. While adding the new features to the statecharts, the refinement rules for B-action systems should be applied. For each refinement step a new set of class and statechart diagrams are generated. The refined diagrams are then automatically translated to B-action systems with the U2B-tool.

Currently the U2B tool provides the following support for refinement.

a) The B component is created as a refinement instead of a machine (*i.e.*, it has a `REFIMEMENT` header and a `REFINES` clause).

b) Refinement of events can be indicated on statecharts by merging and splitting transitions on statecharts.

5.6.1 Refinement in Class Diagrams

During the development more implementation details are added to the system. The new features are modelled with new attributes. Sometimes even new operations have to be added in the class diagrams. In Figure 5.9 the attributes *awaited*, *move_counter* and *analyse_counter* have been added to the Analyser class to keep track of the services performed by the Analyser. In the class *A_PROC* the new features on the operating table of the Analyser are given. The attribute *zpos* gives the height of the operating table in the Analyser and *blank* states whether there is a plate on the operating table in the Analyser or not. Furthermore, new failure-operations are added to the methods of the Analyser.

We state the services of the basic statechart with more details in a refined statechart diagram. This is shown in Figure 5.10 where the component can evolve from a state *Idle* to a state which performs each of the required services, provided certain conditions, *ServiceN_prec*, are satisfied. If these services are successfully performed, when *ServiceN_postc* holds, the component returns to its state *Idle*. An exception of a service suspends the autonomous behaviour of a component, that evolves to the state *Suspension*. If a remedy for an occurred exception is found, the component can resume its service-providing state. If no remedy exists, then the exception is a failure and the component aborts its execution.

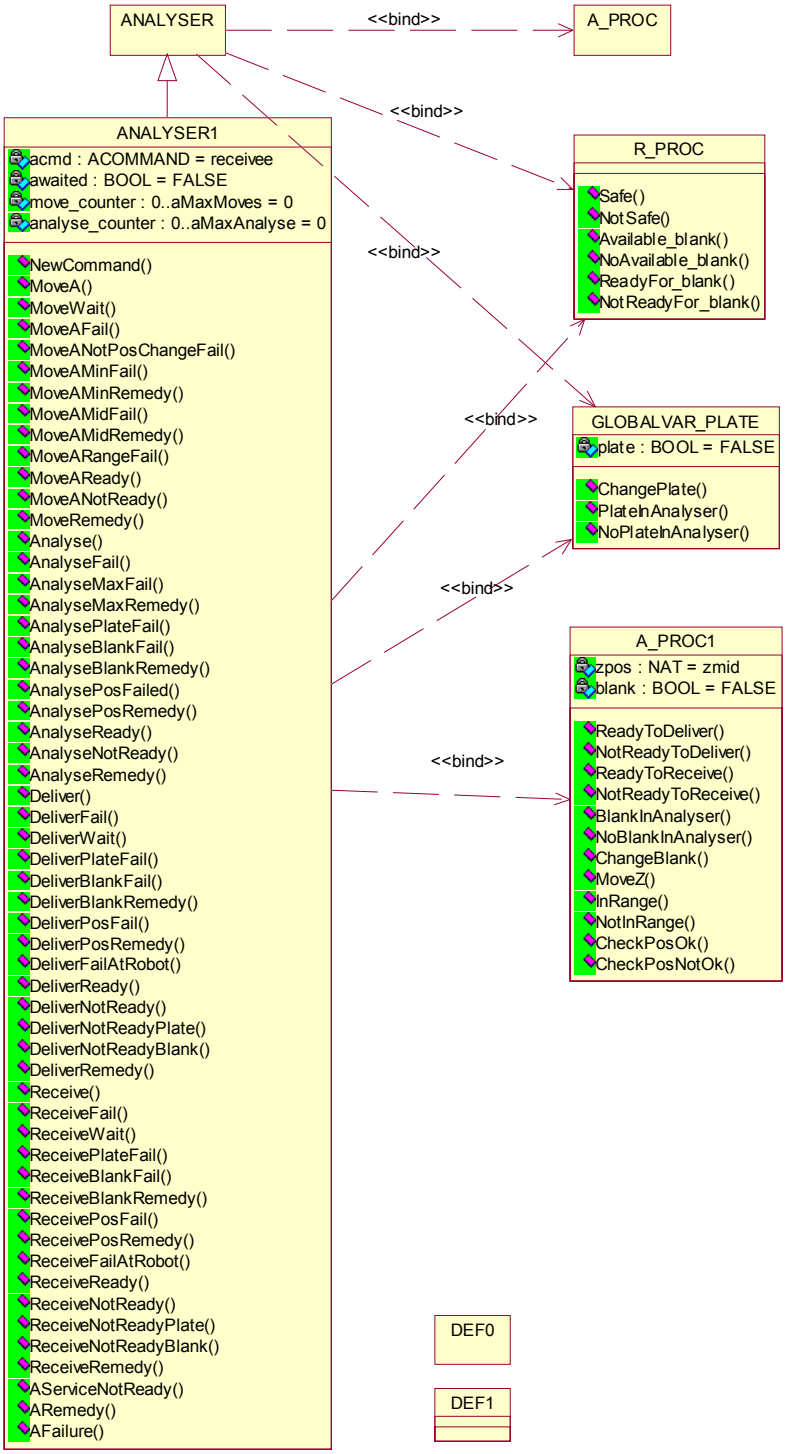


Figure 5.9: A refined class diagram of the Analyser.

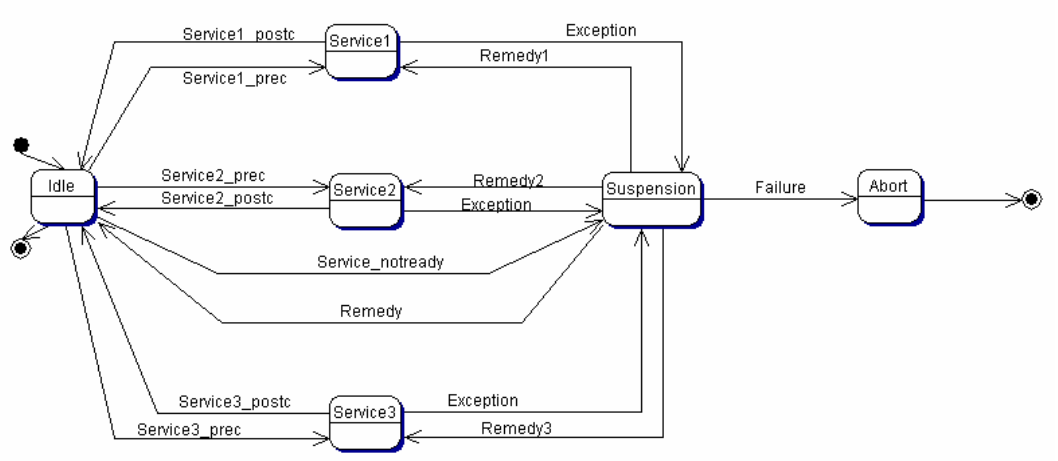


Figure 5.10: A refined statechart diagram for a component.

The refined statechart diagram is derived from the use case description (subroutine description) and the refined class diagram. All the conditions shown in the diagram are expressions on the class attributes and methods. At this level, the initial informal specification regarding the respective component is completely captured by this refined statechart diagram. This diagram can still be refined to a more detailed statechart diagram involving more attributes and methods. The class and statechart diagrams of each refinement step are translated to B. Figure 5.11 and Figure 5.11 show part of the refined statechart diagram of the Analyser. Note that similar diagrams for the *move* and *deliver* commands and the *service_not_ready* condition exist, but are not shown.

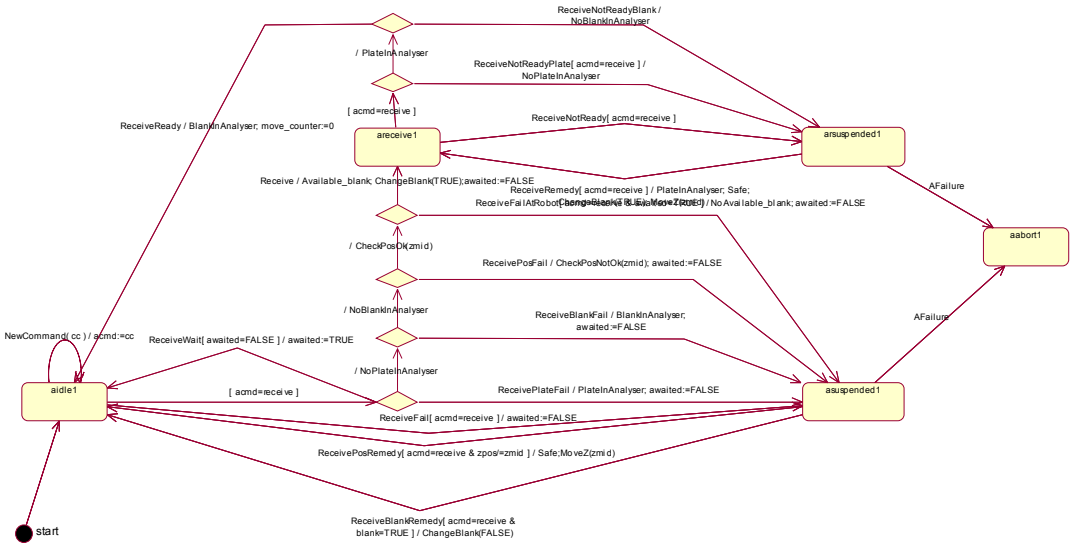


Figure 5.11: Transitions associated with a receive command.

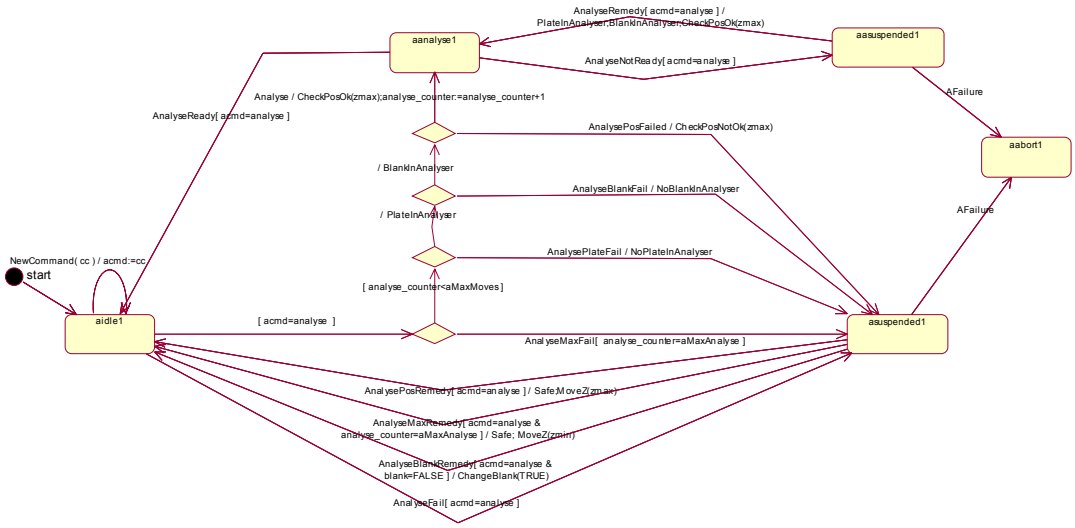


Figure 5.12: Transitions associated with an analyse command.

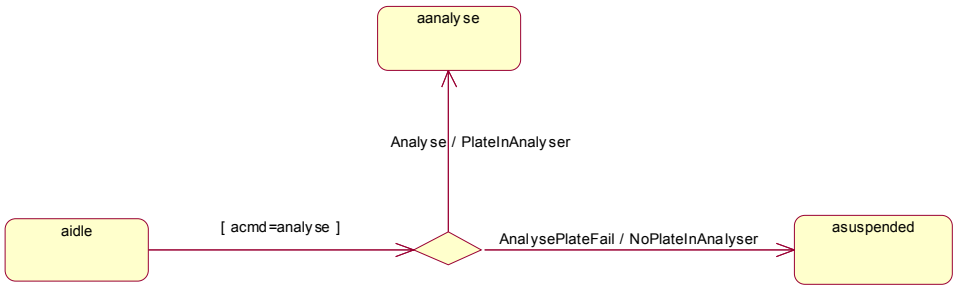
5.6.2 Refinement using Statecharts

The rest of the section describes in detail how both transitions and states can be involved in the refinement.

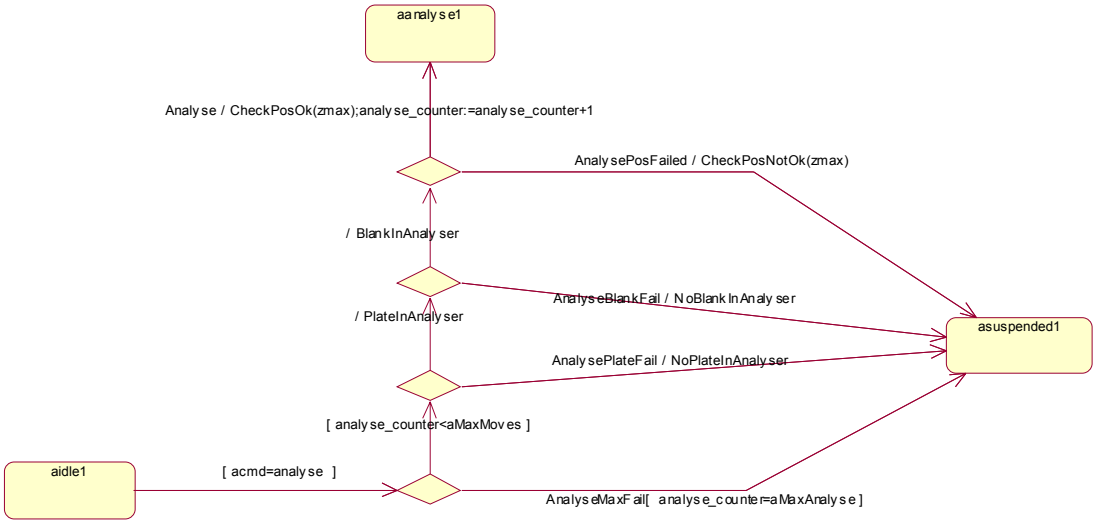
5.6.2.1 Transition Refinement

One technique that was employed was to break down a single transition, representing one event, into several different transitions representing alternative events that are either refinements of the original event or additional events (refining the non-event, `skip`).

For example, in the analyser example the state transitions associated with a command to analyse were



These were refined by adding new transitions representing additional failure conditions. In doing so the transition path from *aidle* to *aanalyse* may be refined by additional guards and actions.



The corresponding fragments of B (produced by U2B) are:

For the abstract actions of ANALYSER:

```
Analyse  =
   BEGIN
     SELECT 0=0
     THEN astate:=aanalyse ||
          PlateInAnalyser ||
          SELECT astate=aidle ∧ acmd=analyse THEN skip END
     END
   END
;
AnalysePlateFail  =
   BEGIN
     SELECT 0=0
     THEN astate:=asuspended ||
          NoPlateInAnalyser ||
          SELECT astate=aidle ∧ acmd=analyse THEN skip END
     END
   END
```

For the refined actions of ANALYSER1:

```
Analyse  =
    BEGIN
       SELECT 0=0
       THEN   astate:=aanalyse1 ||
           CheckPosOk(zmax);analyse_counter:=analyse_counter+1 ||
           SELECT 0=0
           THEN   BlankInAnalyser ||
               SELECT 0=0 THEN  PlateInAnalyser ||
                   SELECT analyse_counter < aMaxMoves
                   THEN  SELECT astate=aidle1 ∧ acmd=analyse THENskip END
                   END
               END
           END
       END
```

```
    END;
AnalyseMaxFail  =
    BEGIN
        SELECT  analyse_counter=aMaxAnalyse
        THEN   astate:=asuspended1 ||
               SELECT astate=aidle1 ∧ acmd=analyse THEN skip END
        END
    END;
AnalysePlateFail  =
    BEGIN
        SELECT 0=0 THEN  astate:=asuspended1 ||
           NoPlateInAnalyser ||
           SELECT analyse_counter<aMaxMoves
           THEN  SELECT astate=aidle1 ∧ acmd=analyse THEN skip END
           END
        END
    END;
AnalyseBlankFail  =
    BEGIN
        SELECT 0=0
        THEN   astate:=asuspended1 ||
           NoBlankInAnalyser ||
           SELECT 0=0
           THEN   PlateInAnalyser ||
              SELECT analyse_counter<aMaxMoves
              THEN  SELECT astate=aidle1 ∧ acmd=analyse THEN skip END
              END
           END
        END
    END;
AnalysePosFailed  =
    BEGIN
        SELECT 0=0 THEN  astate:=asuspended1 ||
           CheckPosNotOk(zmax) ||
           SELECT 0=0 THEN  BlankInAnalyser ||
              SELECT 0=0
              THEN   PlateInAnalyser ||
                 SELECT analyse_counter<aMaxMoves
                 THEN  SELECT astate=aidle1 ∧ acmd=analyse THEN skip END
                 END
              END
           END
        END
    END
```

5.6.2.2 State Refinement

Another refinement that was used on the statechart models was to break down a single state into several sub-states. Currently, the U2B translation does not support hierarchical states. Therefore, we refined the original set of states with a complete new set of states replacing the refined state with its sub-states. This involves renaming all the states, even those that are not being decomposed. For example, the *asuspended* state of *ANALYSER* was broken down into different sub-states in order to facilitate enabling of the correct remedy action for the refined failures. This meant that, in the translated B, the state variable was refined and glued as shown below:

```
MACHINE ANALYSER
…
SETS
     ASTATE={aidle,amove,aanalyse,areceive,adeliver,aabort,asuspended}
VARIABLES
     astate, …
INVARIANT
     astate : ASTATE & …
INITIALISATION
     astate := aidle || …
```

is refined by:

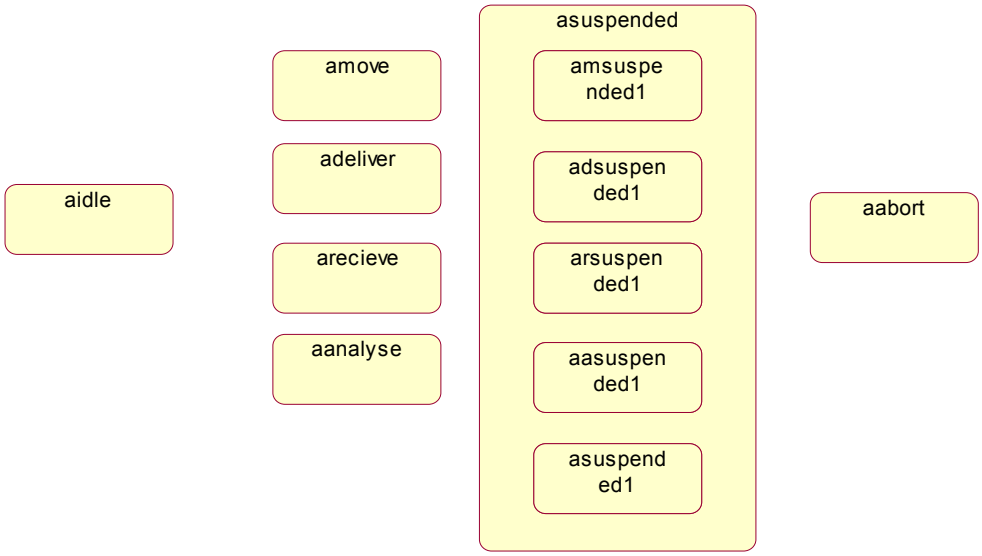
```
REFINEMENT ANALYSER1
REFINES
     ANALYSER
…
SETS
     ASTATE={aidle1,amove1,aanalyse1,areceive1,adeliver1,aabort1,asuspended1,
             amsuspended1,aasuspended1,adsuspended1,arsuspended1}
VARIABLES
     astate,…
INVARIANT
     astate ∈ ASTATE ∧
…
     (astate1=aidle1 ⇒ astate=aidle) ∧
     (astate1=amove1 ⇒ astate=amove) ∧
     (astate1=aanalyse1 ⇒ astate=aanalyse) ∧
     (astate1=adeliver1 ⇒ astate=adeliver) ∧
     (astate1=areceive1 ⇒ astate=areceive) ∧
     (astate1=asuspended1 ⇒ astate=asuspended) ∧
     (astate1=amsuspended1 ⇒ astate=asuspended) ∧
     (astate1=aasuspended1 ⇒ astate=asuspended) ∧
     (astate1=adsuspended1 ⇒ astate=asuspended) ∧
     (astate1=arsuspended1 ⇒ astate=asuspended)
INITIALISATION
     astate := aidle1 ||
…
```

5.6.2.3 Hierarchical State Refinement

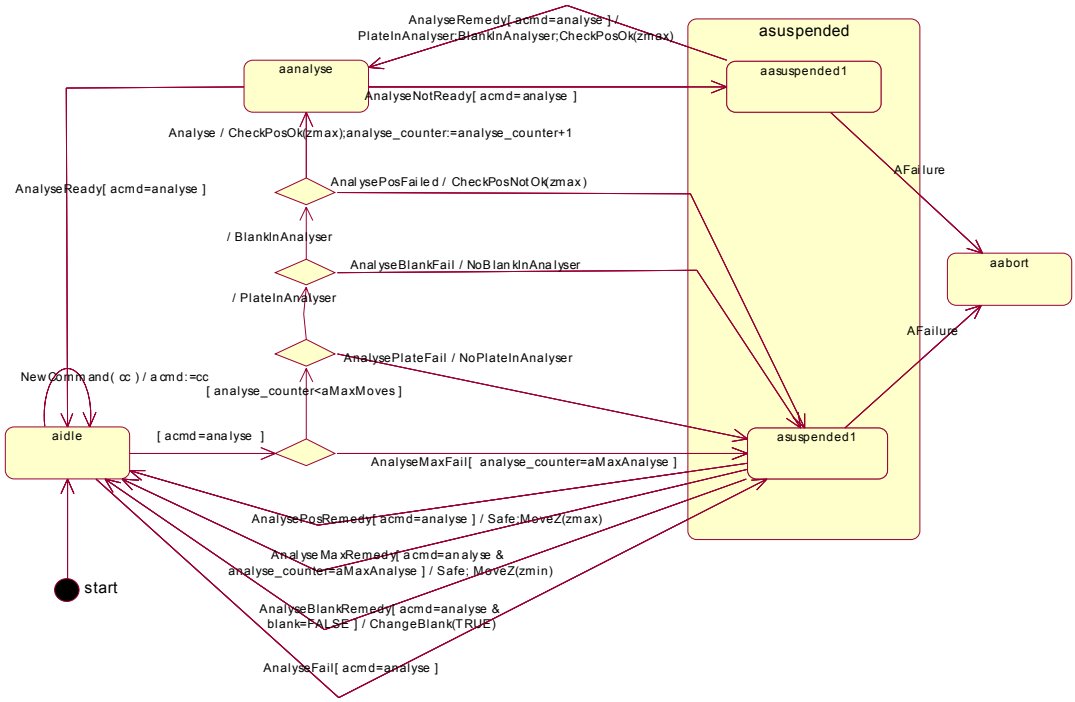
As observed above, the correspondence of all the states has to be explicitly stated in the invariant even when there is no real change to the state (*e.g.*, `aidle1` is equivalent to `aidle`). An improvement to the U2B translator, so that it handles hierarchical states would enable the implicit refinement correspondence of B to be utilised and the hierarchical correspondence would become more obvious. The resulting superposition refinement would also make the proofs easier to discharge. The proposed new method of handling state refinement is shown below.

This diagram shows the states with no transitions. It illustrates the hierarchical relationship between the states in the statechart model of *ANALYSER1_H* (a version of *ANALYSER1* utilising this proposed method):

Note that, apart from the addition of sub-states of *asuspended*, the set of states are exactly the same as in *ANALYSER* (the class refined by *ANALYSER1_H*) in contrast to previous refinements where all the states were renamed by post fixing the number 1.



The statechart diagrams of *ANALYSER1_H* look exactly the same as those of *ANALYSER1*, except that they use the states shown in this diagram. That is, they use the state from *ANALYSER* wherever the state is not decomposed, and the relevant sub-state wherever it is. The parent state, *asuspended*, is not used on any statechart diagram, it is implied by the transition entering one of its sub-states. For example, the (partial) analyser statechart shown in Figure 3.7 becomes:



Note that showing the parent state, *asuspended*, on the diagram is superfluous. This is because the hierarchy is captured in the model, not on the diagram. The B refinement that we would like the U2B translator to produced is shown below.

Note:

1. Only operations relevant to the analyser command are shown.
2. the gluing invariant is generated automatically from the state hierarchy whereas before it had to be manually entered as information attached to the class.

```
REFINEMENT ANALYSER1
REFINES
     ANALYSER
SEES
     DEF0, DEF1
INCLUDES
     R_PROC,
     GLOBALVAR_PLATE,
     A_PROC1
SETS
     ASTATE={aidle,amove,aanalyse,arecieve,adeliver,aabort,asuspended}
     ASUSPENDED_SUBSTATE={asuspended1,amsuspended1,aasuspended1,adsuspended1,
                          arsuspended1}
VARIABLES
     astate,
     asuspended_substate,
     acmd,
     awaited,
     move_counter,
     analyse_counter
INVARIANT
     astate ∈ ASTATE ∧
     asuspended_substate ∈ ASUSPENDED_SUBSTATE ∧
     acmd ∈ ACOMMAND ∧
     awaited ∈ BOOL ∧
     move_counter ∈ 0..aMaxMoves ∧
     analyse_counter ∈ 0..aMaxAnalyse ∧
     (asuspended_substate=asuspended1 ⇒ astate=asuspended) ∧
     (asuspended_substate=amsuspended1 ⇒ astate=asuspended) ∧
     (asuspended_substate=aasuspended1 ⇒ astate=asuspended) ∧
     (asuspended_substate=adsuspended1 ⇒ astate=asuspended) ∧
     (asuspended_substate=arsuspended1 ⇒ astate=asuspended)
INITIALISATION
     astate := aidle ||
     asuspended_substate :∈ ASUSPENDED_SUBSTATE ||
     acmd := receivee ||
     awaited := FALSE ||
     move_counter := 0 ||
     analyse_counter := 0

OPERATIONS
AnalyseFail=
     BEGIN
          SELECT    astate=aidle ∧ acmd=analyse
          THEN astate:=asuspended || asuspended_substate:=asuspended1
          END
     END;
AnalyseReady=
     BEGIN
          SELECT    astate=aanalyse ∧ acmd=analyse
```

```
          THEN astate:=aidle
          END
     END;
AnalyseNotReady=
     BEGIN
          SELECT    astate=aanalyse ∧ acmd=analyse
          THEN astate:=asuspended || asuspended_substate:=aasuspended1
          END
     END;
AnalysePosRemedy=
     BEGIN
          SELECT    astate=asuspended ∧ asuspended_substate=asuspended1 ∧
                    acmd=analyse
          THEN astate:=aidle ||
               Safe;MoveZ(zmax)
          END
     END;
AnalyseMaxRemedy=
     BEGIN
          SELECT    astate=asuspended ∧ asuspended_substate=asuspended1 ∧
                    acmd=analyse ∧ analyse_counter=aMaxAnalyse
          THEN astate:=aidle ||
               Safe; MoveZ(zmin)
          END
     END;
AnalyseBlankRemedy=
     BEGIN
          SELECT    astate=asuspended ∧ asuspended_substate=asuspended1 ∧
                    acmd=analyse ∧ blank=FALSE
          THEN astate:=aidle ||
               ChangeBlank(TRUE)
          END
     END;
AnalyseRemedy=
     BEGIN
          SELECT    astate=asuspended ∧ asuspended_substate=aasuspended1 ∧
                    acmd=analyse
          THEN astate:=aanalyse ||
               PlateInAnalyser;BlankInAnalyser;CheckPosOk(zmax)
          END
     END;
AnalyseMaxFail=
     BEGIN
          SELECT     analyse_counter=aMaxAnalyse
          THEN astate := asuspended || asuspended_substate:=asuspended1 ||
               SELECT astate = aidle ∧ acmd=analyse THEN skip END
          END
     END;
AnalysePlateFail=
     BEGIN
          SELECT    0=0
          THEN astate := asuspended || asuspended_substate:=asuspended1 ||
               NoPlateInAnalyser ||
               SELECT    analyse_counter<aMaxMoves
               THEN
                    SELECT astate = aidle ∧ acmd=analyse THEN skip END
               END
          END
     END;
AnalyseBlankFail=
     BEGIN
          SELECT    0=0
          THEN astate := asuspended || asuspended_substate:=asuspended1 ||
               NoBlankInAnalyser ||
               SELECT    0=0
               THEN PlateInAnalyser ||
```

```
                    SELECT    analyse_counter<aMaxMoves
                    THEN
                         SELECT astate = aidle ∧ acmd=analyse THEN skip END
                    END
               END
          END
     END;
Analyse=
     BEGIN
          SELECT    0=0
          THEN astate := aanalyse ||
               CheckPosOk(zmax);analyse_counter:=analyse_counter+1 ||
               SELECT    0=0
               THEN BlankInAnalyser ||
                    SELECT    0=0
                    THEN PlateInAnalyser ||
                         SELECT    analyse_counter<aMaxMoves
                         THEN
                            SELECT astate = aidle ∧ acmd=analyse THEN skip END
                         END
                    END
               END
          END
     END;
AnalysePosFailed=
     BEGIN
          SELECT    0=0
          THEN astate := asuspended || asuspended_substate:=asuspended1 ||
               CheckPosNotOk(zmax) ||
               SELECT    0=0
               THEN BlankInAnalyser ||
                    SELECT    0=0
                    THEN PlateInAnalyser ||
                         SELECT    analyse_counter<aMaxMoves
                         THEN
                           SELECT astate = aidle ∧ acmd=analyse THEN skip END
                         END
                    END
               END
          END
     END

END
```

6 Implementing Models

This chapter provides guidelines on automatically generating code from B implementations and integrating this with other code.

6.1 Producing Efficient Code Automatically

One of the main advantages of using the B method is automatic code generation. In fact, at the end of the refinement process, the final component is an implementation in a subset of the B language, B0. One of the main questions when starting this formal development was about efficient code generation. We have chosen to develop a simple code translator. The idea is to use it as a prototype to figure out what kind of improvements can be implemented and what kind of improvements are necessary. Note that improvement can be dedicated to a single smart card chip target but we do not consider this point here.

Our translation principles are similar to the conversion scheme of the Atelier-B Tool, with some minor modifications to reduce memory consumption. For example, the constant values are converted to pre-processor directives, instead of variables. Each implemented machine is converted into a C source file and a corresponding C header file. Although only the code contained within the implementation machine generates corresponding C source code, the conversion process needs to take into account the abstract machine and all its refinement steps. For example, concrete variables can be introduced at every refinement step. However, the signature of the operations is fully defined in the abstract machine: this is a B requirement, since an implementation that imports a machine can call operations defined in that machine, but has no knowledge of the implementation details.

It should be noted, however that a user of the converter only needs to provide the implementation to the converter: the implementation machine contains enough information for the converter to retrieve the refinements and the abstract machine.

The header file defines the enumerated sets definitions, the constants definitions and the C functions prototypes. The C source file contains the actual C conversion of the machine's operations. Although those files are generated for most machines, not every machine is converted to a C source file. For example, context machines that only define constants or enumerated sets are translated into a C header file defining the corresponding constants or enumeration.

6.1.1 Machine Conversion

Although the B0 language is similar to classical imperative languages, an important point when converting machines is that a system may contain multiple instances of a given machine. This makes converting B0 to C similar to converting an object oriented language to C. To allow the use of multiple machine instances, the state of the converted machine is defined within a C structure. The state consists of all the concrete variables, and of the imported machines' state.

To allow operations to modify a specific machine instance, a parameter corresponding to a pointer to the instance is added to each operation. This pointer is similar to the `this` pointer used by C++ or Java.

Enumerated sets are converted to C `enums`, and the pre-processor is used to define constant values.

Architecture clauses such as `IMPORTS` and `SEES` are all translated using the include directive of the C pre-processor: although this directive does not enforce the access rules of the B clauses, as the B specification complies with those rules, so will the generated code.

Finally, the initialisation clause is treated as a special operation that performs the initialisation of the variables and calls the initialisation function of the imported machines. This function has to be called before any other functions are invoked.

6.1.2 Operation Conversion

Converting operations is quite straightforward, since the constructs allowed by B0 match the constructs allowed in C. The conversion involves replacing the B substitutions by the equivalent C instructions and the variables by their C equivalents.

Whereas B operations may have several return parameters, C functions have only one. When an operation has more than one return parameter, the first parameter is returned as the C function return value, and the next return values are passed by reference as pointers to the destination variables. This increases the converter complexity, since the converter has to know how a given variable will be returned. Moreover, for return values passed by reference, the converter has to differentiate occurrences of those variables appearing on left or right sides of substitutions.

Finally, in order to prevent name clashes, all operations names are prefixed with the name of the machine.

A converter prototype has been developed using Java and the JavaCC B parser distributed on [Tatibouet99]. The aim of this converter is mainly to test the feasibility of the approaches considered and it currently only handles a subset of the B language.

Another limitation of this converter is that the verification of the correctness of the optimisations performed has not been studied at this prototype stage. A translator meant to be used in industrial development should be developed so that it could be certified, for example by a Common Criteria evaluation.

6.2 Timing Constraints in Code Generation

The objective is a cycle time of the executable code, running on target, less or equal than 330 ms.
First of all, the time performance can depend on the controller, but not on the physical part. And the controller is only a specification, the implementation is made at software level, so the time optimization (if required) is also made at a low software level.
A conclusion of the railway case study is that the controller has very little influence on timing performances of the software, which depends quite entirely on the choices made during the implementation of the software.

6.3 Resource Constraints in Code Generation

This section describes some optimisations that can be added to the previous translation scheme in order to reduce the memory consumption of the generated code. The objective of those optimisations is not to replace the optimisations performed by the C compiler, but to address optimisations that the compiler cannot perform on its own, since the converted code would not contain enough information.

Most of these optimisations will require the translation to be performed on a project basis. That is, a machine cannot be converted without knowing how the other machines are converted. The new conversion process is shown in Figure 6.1.

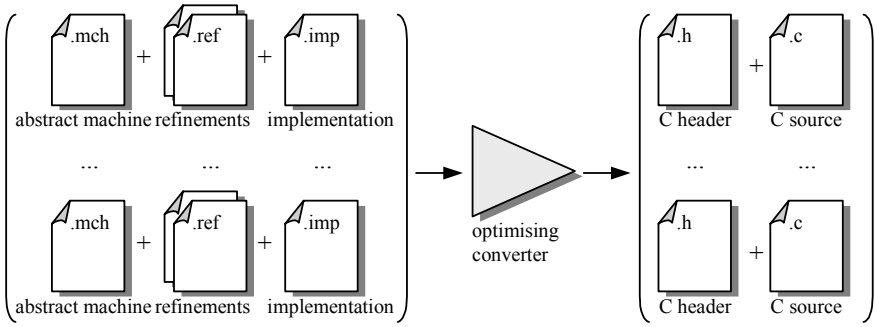


Figure 6.1: Optimised conversion of a B project.

6.3.1 Smart Card Constraints

Smart cards are devices with severe constraints both for memory size and computing power (The chips used are usually 8-bit chips, and smart card memory rarely exceeds 64 kilobytes of ROM, 64 kilobytes of EEPROM, and 4 kilobytes of RAM). So, although formal methods have shown their efficiency for proof and validation of specification, the code generated from those formal specifications can usually not be embedded within a smart card. This requires either limiting the use of formal methods to specification, or manually translating the specification.

The B method has already been used to specify some smart card operating system components. Although the code automatically generated from B specifications cannot be embedded yet, it is very close to meet smart cards constraints, and could meet them if an optimising code generator is used.

6.3.2 Machine Instances

When a machine has no state on its own, it is possible to suppress the use of the additional *instance* parameter. We consider that a machine has no state when it does not define any concrete variable, and imports only machines that have no state.

An example of machine without state would be a machine encapsulating an algorithm. Figure 6.2 shows a simple example of such a machine, and its corresponding implementation.

```
MACHINE  Average
OPERATIONS
res ← average(val1, val2)
=
PRE
    val1 ∈ NAT ∧
    val2 ∈ NAT ∧
    val1 + val2 ∈ NAT
 THEN
    res := (val1 + val2) /
2
 END
END
```

```
IMPLEMENTATION  Average_i
REFINES
    Average
OPERATIONS
 res ← average(val1,
val2) =
 BEGIN
    res := (val1 + val2) /
2
 END
END
```

Figure 6.2: A machine encapsulating an algorithm.

A conversion of this machine removing the instance parameter is shown in Figure 6.3. The same strategy can be used to translate machines that are instantiated only once within the project to convert: in this case the pointer used to access multiple instances can be suppressed, since its value will always be the same, and known statically.

```
int average(int val1, int val2) {
    int res;
    res = (val1+val2)/2;
    return res;
}
```

Figure 6.3: C Translation of the `Average` machine.

It should be noted that, as the operations calling convention changes depending on the way the machines are translated, such an optimisation could not be performed to translate machines without knowledge on how the other machines are converted.

6.3.3 Variables Storage

A simple conversion from B0 to C translates all the integer values into C *int*. However, it is possible to use smaller integers depending on the range used by the B variables. This allows using the *short* and *char* C types to store the variables, and can greatly decrease the memory consumption. For smart cards applications, which mainly use *byte* and *short* values, the memory required can be divided by more than two if the adequate type is used.

```
MACHINE
    Range
…
INVARIANT
    begin ∈ 0..255 ∧ end ∈ 0..255 ∧
    begin ≤ end
...
END
```

Figure 6.4: Machine declaring *byte* variables.

For example, assuming that the variables of the `Range` machine are specified as *byte* intervals (Figure 6.4), the converter could use this information to assign a corresponding C type.

A possible C translation of the *getRange* operation is provided in Figure 6.5. It assumes that the variables *begin* and *end* have been typed as *char*.

```
unsigned char getRange(unsigned char
                       *r_max)
{
    unsigned char r_min;
    r_min  = begin;
    *r_max = end;
    return r_min;
}
```

Figure 6.5: C translation of the *getRange* operation.

Using different C types for storing variables can however introduce memory corruption if not handled carefully: as some results are passed by reference using C pointers, the types of the variables have to be the same.

```
MACHINE
    Range_User
INCLUDES
    Range
CONCRETE_VARIABLES
    var1, var2
INVARIANT
    var1 ∈ NAT ∧
    var2 ∈ NAT ∧
OPERATIONS
    useRange = BEGIN
     var1, var2 ← getRange
    END
END
```

```
IMPLEMENTATION
    Range_User_i
REFINES
    Range_User
IMPORTS
    Range
...
INITIALISATION
    var1 := 0;
    var2 := 0
OPERATIONS
    useRange = BEGIN
     var1, var2 ← getRange
    END
END
```

Figure 6.6: Implementation importing the `Range` machine.

The `Range_User` machine presented in Figure 6.6 is a valid B machine that needs a special conversion: as the variables `var1` and `var2` are natural numbers, the results of the `getRange` operation can safely be stored inside `var1` and `var2` without overflow. However, assuming that the variables `var1` and `var2` are allocated as integer variables, and that the `getRange` operation is translated as shown in Figure 6.5, passing the address of the variable `var2` to the `getRange` operation would only update one byte of this variable, leading to erroneous results.

As shown in Figure 6.7, a workaround is to use a temporary variable to pass the parameter when calling the operation. The overhead of using temporary variables can however balance the savings obtained by using smaller types if the specification is not written with this concern in mind. So, always allocating the smallest type for every variable may not be a good choice. A better solution would be to perform an analysis of the complete machine set to be translated, and to allocate variable types depending on the number of conversions that will be needed.

```
int var1, var2;
void useRange()
{
```

```
    unsigned char tmp_var2;
    var1 = getRange(&tmp_var2);
    var2 = tmp_var2;
}
```

Figure 6.7: Safe conversion of the *useRange* operation.

6.3.4 Initialisation Translation

Initialisation can often be performed statically instead of using a dedicated initialisation function. This solution allows suppressing all the initialisation code, but is more difficult to implement, since this initialisation operation may have to be interpreted by the converter.

6.3.5 Operation Inlining

Inlining an operation call corresponds to replace the corresponding function call by the body of the function.

Operation inlining can be used basically for two objectives. The first one corresponds to the classical use of in-lining. That is inlining simple functions where the inlined code is cheaper than a function call. For example, for simple function that only stores a value in a variable, the cost of the function call can be avoided. Moreover, in such cases, the number of assembly instructions generated is likely to be less than the number of instructions that a function call would generate.

The second objective of inlining is to remove operation calls that have not been introduced to factorise code, but for specification purpose: we have found that, in order to simplify the specification, or to ease the proof process, specifications are structured using multiple machines and operations.

Although this strong decomposition is usually a good specification practice, it also involves a large number of operation calls. Those calls are costly both in terms of computing resources, and of memory consumption.

Inlining those operations allows the use of a clear and modular specification, without the overhead of the generated function calls. To achieve this, a simple inlining criterion would be to inline all the operations that are called from only one location within the specification.

6.4 Integrating Formal and Non-Formal Code

In the smart case study the chip target is an ATMEL platform. First, we provide an implementation on a ATMEL AT90 SC 3232. This chip contains 32 kb for the software, 32 kb for the data and 1.5 kb of RAM. This implementation allows us to demonstrate the feasibility of embedding a formally developed verifier into a smart card. This first implementation embeds a complete type verifier except the *integer* management, and a partial structural verifier where external tests are excluded. This first implementation required 29 kb of compiled code and fits in the ATMEL chip. The code of the verifier is stored in the first 32 kb for the software. Applets are loaded into the 32 kb dedicated to data.

The second implementation we provide is on an ATMEL AT90 6464 C with 64 kb for software, 64 kb for data and 3 kb of RAM. On this platform, we aim to embed the complete verifier, with the `integer` management for the type verifier and all the structural tests. With no specific optimisation, the size of this complete verifier is 45 kb.

The compilation chain provided by ATMEL, which has some very efficient tools such as the compiler, which enable significant space (in terms of code size) savings to be made. Among the different options available for compiling the code for ATMEL, we chose to focus on code size and not yet for code execution efficiency. Our priority is to restrict code size over runtime efficiency, in order to embed the byte code verifier within a smart card.

6.5 Integrating Testing and Formal Development

Software development requires a description of the requirements. These requirements generally do not describe the solution, but the needs by using a functional approach. This description mainly concerns the inputs and outputs of the system, and the expected results in some specific circumstances. One of the project leader responsibilities is to define a specification of the solution, and to validate that this solution corresponds to the needs informally expressed. A formal development can then be used, in order to ensure that the product is compatible with its specification.

Considering the Java Card verifier, the verifier requirements are expressed through the Java Card specification, but no verifier specification is given. The first step when developing a Java Card verifier is to write specifications that are supposed to satisfy explicit and implicit needs. Then, a verifier can be developed using a formal development process. The overall process is illustrated in Figure 6.8.

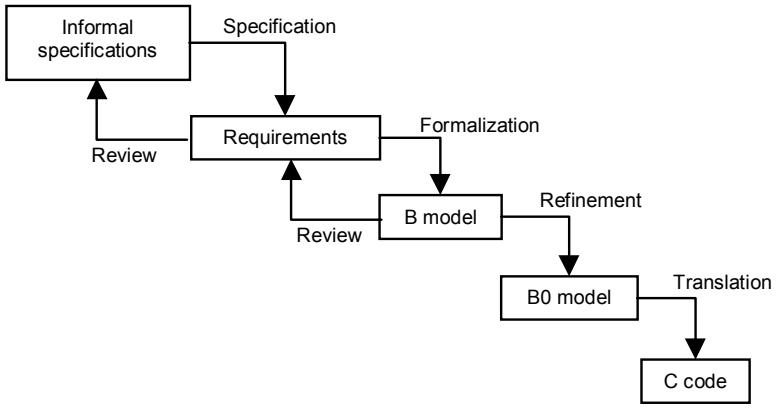


Figure 6.8: Formal development process including early steps.

Two major problems can still arise, that are very time and money consuming:

- The requirements can often be ambiguous or incomplete, and an informal validation may not identify this. Thus, although the resulting software conforms to its specification, it will not meet its requirements. One solution is to formalise the software environment, in order to limit the risk of ambiguity or incompleteness. It then becomes possible to generate tests

suites based on the environment, and to validate software integration within this environment.

- The execution platform does not conform to its specification, generating failures during execution. The software developed does not provide the expected results because of them. In order to detect these errors, the test process has to include not only functional tests but also robustness tests.

The risk of these problems occurring can be reduced using testing in the manner illustrated in Figure 6.9.

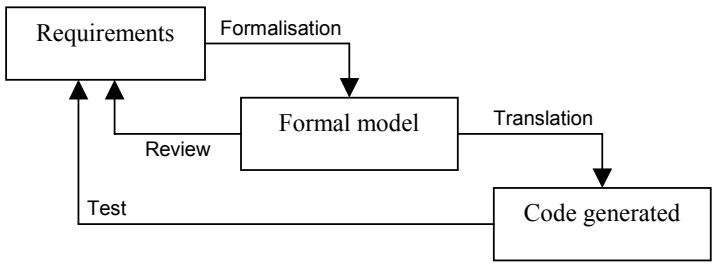


Figure 6.9: Using test in conjunction with a formal development.

6.5.1 Test Generation Automation

Several techniques exist to automate test generation. In a formal development, where a trusted translator automatically generates the code, and when the code is executed on a trusted platform, no conformance testing is required. The last problem is then to check that the application fulfils its environment's needs. In order to automate these tests, a possible solution is to formalise a model of the environment and to generate tests that check for conformity to the environment behaviour. When the trusted formal chain is incomplete, conformity tests are relevant; as they limit the risk of errors which are generated by the translator or are present in the execution platforms. The process consists in using the formal specifications of the application to generate tests.

Different types of tests can be generated, depending on the kind of application. Applications that are not critical can be tested using only a few tests. One strategy can be to cover each function, each state or each transition of the model only one time. For more critical applications, tests can cover all, or particular, paths in the model.

In any case, the test strategy is strongly dependent on the client needs in term of quality and security. The major difficulty consists in defining the end of the test process, considering that exhaustive test is almost impossible.

6.5.1.1 Example of Flaws

We considered that most of errors have been identified and eliminated after the proof process. If the refinement process ensures that the implementation conforms its specification, it does not ensure that the corresponding formal specification is correct. In fact, when translating informal

specification into formal ones, errors may have been introduced. This is due to the misunderstanding of the informal specification by the developer. For this reason, tests are still necessary to make sure that these kinds of errors do not exist. Moreover, as the B code is automatically translated into C code by a tool and that this tool is not "qualified", it is also necessary to test the C code in order to check that it is correctly generated. These are the two reasons why testing is still necessary. 14 errors have been discovered during the testing step on the type verifier. It is important to note that these errors cannot be discovered by the proof as it concerns the higher specification. Half of errors concern the model itself. That is, a wrong writing of the formal model. 8 errors concern directly the model of the byte code instructions. These errors are due to the misunderstanding of the informal specification provided to the formal developer. It concerns for example the `invokeinterface`, the `invokespecial` or the `athrow` instructions that are not really easy to understand and thus to model. One may not that for example some errors have already been found on instruction such as the `invokespecial` and the `checkcast` during the code review and the proof activity. That enforces the fact that these instructions are really difficult to model and to implement. The proof activity has removed some of implementation errors and the testing phase helps identifying the specification errors.

6.5.1.2 Testing a Verifier: A Manual Test Strategy

An automated test generation process, requires inputs of the application under test to be automatically generated. For the Java Card verifier, the system under test is shown in Figure 6.10.

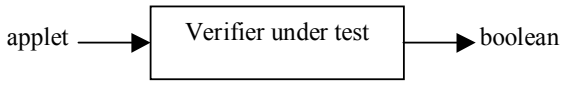


Figure 6.10: Black box testing of a verifier.

In order to automatically generate test suites, the test generation process has to automatically generate different type of test applets. These applets have to be expressed in a specific format in order to be checked by the verifier. In its actual version, this format seems to be too complex to enable any kind of test applet generation (including invalid applets). Another solution could be to generate abstract applets corresponding to concrete applets. Eliciting the different type of applets, valid and invalid, is a process similar to partition analysis but on a very much more complex data type. This could be pertinent if the test process was dedicated to verifier behaviour, with only a few kinds of applets. However, it is exactly the contrary. The verifier is supposed to have no state, focusing on the data only (which in this case is the applet). For all these reasons, the test generation process used in this situation has been manual.

Observe that when the first applet elicitation is performed, it becomes possible to generate automated tests for a verifier, using all the elicited applets as an equivalent of partition analysis.

7 Emerging Complementary Developments

7.1 Event Based Model Decomposition

7.1.1 Introduction

The process of developing an event model by successive refinement steps usually starts with a very few events (sometimes even a single event) dealing with a very few state variables. It usually ends up with many events and many variables. This is because one of the most important mechanisms of this approach consists of introducing new events during refinement steps. The refinement mechanism is also used at the same time to significantly enlarge the number of state variables.

The new events, let us recall, are manifestation of the refinement of the time grain within which we might, more and more accurately, observe and analyse our dynamic system. At some point, we might have so many events and so many state variables that the refinement process might become quite heavy. And we may also figure out that the refinement steps we are trying to undertake are not involving any more the totality of our system (as was the case at the beginning of the development): only a few variables and events are concerned, the others only playing a passive, but noisy, rôle. The idea of model decomposition is thus clearly very attractive: it consists of cutting a heavy event system into smaller pieces which can be handled more comfortably than the whole. More precisely, each piece should be refinable independently of the others. But, of course, the constraint that must be satisfied by this decomposition is that such independently refined pieces could always (in principle) be easily recomposed. This process should then result in a system which could have been obtained directly without the decomposition, which thus appears to be just a kind of "divide-and-conquer" artefact.

This chapter contains the feasibility study of such a mechanism. After proposing an informal definition of decomposition in the next section, we outline the outcome and constraints of this approach (Section 7.1.3). We then present its main difficulty (Section 7.1.4) and propose a solution to it (Section 7.1.5).

7.1.2 Informal Definition

Decomposing an event model M is a complex process which can be defined in the following way:

1. M is first "splitted " (in a certain way to be defined) into several models, say N,… , P.

2. These new models are then be refined independently of each other yielding NR, … , PR.

3. These refined models are eventually "put together" (also in a certain way to be defined) to form yet another model MR.

4. The recomposed model MR is guaranteed to be a refinement of M.

It is important to notice here that point 3 above (the recomposition) will never be performed in practice. One has only to figure out that it can be done under certain circumstances and that the

refinement condition stated in point 4 (MR is a refinement of M) is then satisfied. The events of the component models N, … , P of the decomposition are supposed to form a partition of the events of the initial model M. Concerning the way the variables of M are splitted among N, … , P, the problem is not so simple as we shall see below in section 7.1.4.

7.1.3 Outcome and Constraints of Decomposition

This decomposition process may play three important practical methodological rôles:

1. It is certainly easier (less proofs) to refine (possibly several times) N, … , P independently of each other rather than together.

2. This process of decomposition is "monotonic" in that refinements of N, … , P can be further decomposed in the same way, and so on.

3. The models N, … , P could already possess (off the shelf) some refinements that can then be reused in several projects.

In building this mechanism, we shall observe several "constraints":

1. We must define a process that is totally robust mathematically speaking.

2. We do not want to modify in any way the mathematical definition and concept of refinement.

3. We figure out that the process of decomposition of an event model is distinct (although close) from that of importation as described in "classical B".

7.1.4 The Main Difficulty: Variable Splitting

Suppose that we have a certain model, M, with four events m1, m2, m3, and m4. We would like to decompose M into two separate models:

(1) N dealing with events m1 and m2, and

(2) P dealing with events m3 and m4.

We are interested in doing this decomposition because we "know" that there are some nice refinements that can be performed on m1 and m2 (possibly adding some new events) and independently on m3 and m4 in the same way.

But in doing this event splitting we must also perform a certain variable splitting. Suppose that we have three variables v1, v2 and v3 in M. Like the events, the variables must be splitted too. For instance, we might put v1 and v2 with N (because m1 and m2 are supposedly working with them and not with v3), and thus v3 goes, quite naturally, with P. But the difficulty here is that m3 and m4, which certainly work with v3, might also work with v2, so that, besides v3, P certainly also requires v2 to deal correctly with m3 and m4.

The problem seems unsolvable since apparently it will always exist some shared variables. As a matter of fact, we have the very strong impression that the splitting of the events will always

conflict with that of the variables. Suppose it is not the case. In other words, suppose that, in our example, `m1` and `m2` only work with `v1` and `v2`, while `m4` only does with `v3`. Clearly then, `M` is made of two completely separated groups of events (`m1` and `m2` in one hand, and `m3` and `m4` in the other) which do not communicate in any way with each other. In this case, `M` is obviously made of two distinct models, which could have been handled separately. So, in all interesting cases, the question of shared variables, `v2` in our example, is unavoidable. How are we going to solve this difficulty?

7.1.5 The Solution: Variable Sharing

7.1.5.1 Shared Variables Replication

We have no choice: the shared variables must clearly be replicated in the various components of our decomposition. Notice that the shared variables in question can be modified by any of the components: we do not want to make any "specialisation" of the components, some of them being only allowed to "read", and some other to "write" these variables. We know that it is not possible in general.

The new difficulty that arises immediately at this point concerns the problem of refinement. In principle, each component can freely data-refine its state space. So that the same replicated variable could, in principle, be refined in one way in one component and differently in another: this is obviously not acceptable.

7.1.5.2 A Notion of External Variable

The price to pay in order to solve this difficulty is to give the replicated variables a special status in the components where they stay. Let us call this status: external. An external variable has a simple limitation: it must always be present in the state space of any refinement of the component. In other words, an external variable cannot be data-refined.

7.1.5.3 A Notion of External Event

But this is not sufficient. Suppose that in a certain component an external variable is only read, not written. The trouble with that external variable is that it has suddenly become a constant in that component, which is certainly not what we want. What we need thus in each component, is a number of extra events simulating the way our external variables are handled in the initial model. Such events are called external events. Each of them "mimics", using the external variables only, an event of the initial model that modifies the external variables in question. The reader has understood: "mimic" simply means "is an abstraction of". Of course such external events cannot be refined in their component. Notice that there is a distinction to be made between an external variable and an external event. An external variable is external in all sub-models where it can be found, whereas an external event always has a non-external counterpart elsewhere. An event, however, can be external in several sub-models.

7.1.5.4 Final Recomposition

The recomposition of the initial model by means of refinements of the various components is now extremely simple. We put together all the variables of the individual components

("dereplicating" the various shared variables) and we simply throw away all the external events of each component. It remains now for us to prove that the recomposed model is indeed a refinement of the initial one. Notice again that this recomposition is usually not done explicitly. It is just something that could be done, and which must then yield a refinement of the initial model.

7.2 Vulnerability Analysis

7.2.1 Motivation

The B refinement method produces powerful results; we have a guarantee, through mathematical proof, that the implementation satisfies the original B specification. However, as systems become increasingly complex and distributed it is more likely that mistakes will be made in their design (specification), and that vulnerabilities of such systems will go unnoticed. If a mistake were to be made in the specification of a system, the B development process of refining the specification to code would not identify that mistake.

Consider the *Needham-Schroeder Public Key Authentication Protocol* [Needham78]. The protocol aims to establish mutual authentication between two parties, *initiator* and *responder*, and uses *public key cryptography* in order to do this. The protocol was verified as secure using a variety of techniques, including the *BAN* logic – a logic for reasoning about "belief" of an entity about the system (or context) it is in. But there are multiple attacks upon this protocol [Lowe96], the most well known involving message replay. In the case of the *BAN* logic the rules for reasoning about belief contained some subtle assumptions, which led to a *BAN* model of the protocol system (*i.e.*, the protocol and its context) that was subtly flawed. Hence the proof did not correspond to the protocol as it would operate within its real world context.

If the Needham-Schroeder protocol was written as a B specification (easily achieved since protocols are a form of specification), the process of refining it to code could not have identified the security flaw. It is this limitation on the B method that we address here.

7.2.2 Requirements of the Specification

As a precursor to any B refinement, a technique is needed which verifies that the specification does satisfy the security or safety properties required of it. Such properties are not functional, in that they are specified in terms of what the system should not do, rather than what it should. For example:

"The system should not reveal the contents of the transaction table to anyone other than the administrator",

is a non-functional security property. It is likely that any specification will be required to satisfy a number of such properties simultaneously. However, it is common that such safety or security properties are not defined either formally or informally. Usually such properties are deemed implicit in the functional specifications. This poses a problem for verifying such specifications, since there are no explicit requirements for the specifications to be checked against. In such cases the requirements have to be derived using hazard/vulnerability analysis [O'Halloran99].

The technique described here facilitates the identification of critical properties of systems, and is designed to be used in conjunction with, and to compliment, the B method.

7.2.3 Vulnerability Analysis

The aim is to identify the critical properties of complex systems (*i.e.*, the properties that the corresponding specifications must ensure). Such systems can be viewed as a collection of communicating components. The range of behaviours such systems exhibit is a direct result of the variety of communications that components can enter into, the possible interleaving of such events, and the resulting states of the system. There can be so many different sequences of communications that it is impossible for a human to conceive of every one. This level of system complexity makes it likely that pathological system behaviours or critical properties might be missed, ones which are critical to the correct behaviour of the system. The formalism used to perform such hazard/vulnerability analysis needs to support the debugging, re-programming and refinement of the specification. Further, it needs to support the capture of the non-functional properties that the specification must satisfy.

Tool support for a B development takes the form of theorem proving. Theorem proving tools are unable to provide feedback as to why proofs fail; determining why a proof is impossible can be difficult and time consuming. Further, the B method describes specifications functionally, hence, it is not obvious how B could be used to model non-functional properties. In contrast, model-checkers exhaustively check that every state of an implementation is also a valid state of the specification property. If this is not true, they provide information pertaining to how the system reaches such an invalid state, identifying the trace of events leading up to the property violation. Hence, model checking lends itself naturally to vulnerability analysis.

The process algebra CSP [Hoare85, Roscoe98] allows us to model components in isolation, and then to build models of the total system using the composition of its components. This modular approach to modelling reduces the risk of human error: the user is not required to capture the whole system at once, rather just a small part. It also produces models that are easily traceable to the original specification. Tool support for CSP comes in the form of the model-checker FDR [FDR97]. FDR takes a CSP model of a specification and implementation, and automatically checks that the implementation refines the specification.

Before continuing it is worth mentioning that the success of a vulnerability analysis is dependant on the clarity of the conceptual understanding of what each part of the specification (or implementation) is supposed to do (or does). More subtly, the actual context in which the code is going to be used may have a significant impact on the vulnerability analysis; for example, what interfaces within the system that can be externally monitored, intercepted, or changed may be context dependent. Hence, both a clear layered design of the system and the context in which it is going (intended) to be used are required for a reasonable vulnerability analysis.

Supposing that an adequately clear description of the system and its context are available, and the hazards (or critical information) have been identified, then it is possible to perform various vulnerability analysis techniques [O'Halloran99] to identify if, when, and how a hazard could arise. In the next section a CSP/FDR capability modelling analysis technique is described.

7.2.4 CSP/FDR Capability Analysis

Here it is assumed that the *design,* the system (and context) to be analysed, can be broken down into a collection of capabilities and *inference* rules, which state what capabilities are required to gain new capabilities. Hence, in principle, given an initial set of capabilities it is possible to use these inference rules to determine the potential set of capabilities that can be obtained. If this potential set of capabilities contains a capability representing a *hazard* or *critical information* then it is concluded that either the *hazard* could occur or the *critical information* could be obtained.

At this point it is worth mentioning, that for any *valid* set of capabilities and inference rules, the CSP/FDR vulnerability analysis is fully automatic, returning a tick or a cross depending on whether the analysis was successful. If the analysis failed, then a minimal trace to the *hazard* that could occur, or the *critical information* that could be obtained, is provided for free by FDR's graphical "debug" facility; this provides a list representation the events (inference rules used) to obtain the failure. The minimally of the trace is guaranteed by FDR's breadth first search strategy.

Consider a CSP/FDR design that has a collection of initial capability sets, which represent different entities views of the system. And for each view associate a set of *misuse* capabilities that should not be available from that perspective (*i.e.*, *critical information*). Then the normal expectation would be that the CSP/FDR vulnerability analysis for each view would demonstrate that it was impossible to use the inference rules to obtain any of that view's *misuse* capabilities. Having achieved this, the analyst might want to check how resilient the design was to component failure or capability compromise. This can be achieved by a technique generally known as *fault injection*. Here capability compromise can easily be modelled by adding new capabilities to the initial capability set; and component failures can be modelled by appropriate modification of the inference rules. Further, it is straightforward to construct an automatic check for all *n*-fault designs (*i.e.*, a design with *n* faults). Though, in practice, space and time limits typically limit *n* to somewhere between 1 and 3, depending on the precise nature of the design, and limitations placed on the faults to be injected.

A more technical description of the CSP/FDR analysis is now illustrated by the example in the next section, where the formal CSP statements are written in a `typewriter` font, and the keywords are in **bold**. For a brief overview of the CSP syntax refer to Appendix B; and for further details refer to [Roscoe98,FDR97].

7.2.5 Simplified Smart Card CSP/FDR Capability Analysis Example

7.2.5.1 Interface / Syntax Specification

The simplified example presented here was inspired by the GemPlus smartcard case study. Here we consider the context to be an abstract Java smart card, which contains one purse applet and three loyalty applets, which are uniquely identified by their names. Note that it is useful to have additional no applet identifier (`no_aid`) for the modelling of the applet interfaces.

```
datatype AID = purse | loyalty1 | loyalty2 | spy | no_aid

LoyaltyAID = { loyalty1, loyalty2, spy }
```

In order for the applets to interact they send messages to each other over specified interfaces, namely the purse-to-loyalty interface, the loyalty-to-purse interface, and the loyalty-to-loyalty interface. Given that PurLoyMsg, LoyPurMsg, and LoyLoyMsg represent the messages they are allowed to send on these interfaces respectively, then the following CSP models the interface description, where the first argument represents the source of the message, the second the destination, and the third the content.

```
channel pur_loy_int : AID . AID . PurLoyMsg
channel loy_pur_int : AID . AID . LoyPurMsg
channel loy_loy_int : AID . AID . LoyLoyMsg
```

The following diagram illustrates the model to date.

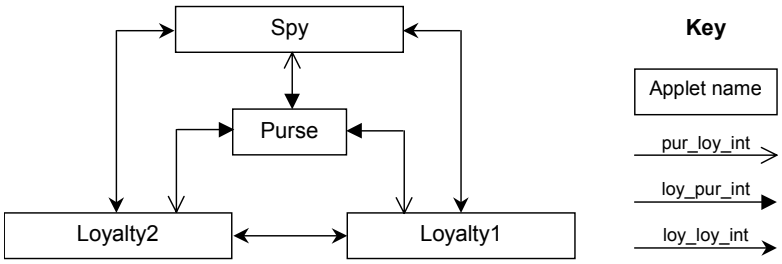


Figure 7.1: Simple CSP case study – structural overview.

The following table summarises the messages, in this case commands and their responses, which can be sent over these interfaces.

Cmd. Name	var:Type	Command Description
`log_full`		The purse sending a command to indicate to the registered loyalties that its transaction database is full.
`get_next_trn`	`l:LoyaltyAID`	A loyalty sending the get next transactions command for loyalty `l`. Note that for the purposes of modelling a transaction can be represented by the loyalty it is associated with.
`get_next_trn_r`	`r:AID`	The get next transaction command returns a result `r`, which is: `none` if either there is no next transaction record for loyalty `l`, or this action is refused for some reason `l` otherwise
`avail_funds`	`l:LoyaltyAID`	A loyalty `l` sending an "available funds" command to another loyalty.
`avail_funds_r`	`r:Set(AID)`	The available funds command returns a result `r`, which is a loyalty-indexed set of funds, where the indexes are used to identify how much each loyalties has contributed. For the purposes of modelling this can be abstracted to the set of applet identifiers that contribute to the funds.

Table 5: Simple CSP case study – message overview.

The information in the above table is modelled formally by the following CSP, where each command is represented by a pair (`c:CmdName`, `v:Value`), where `Value` is either an `AID` (or `LoyaltyAID`), `Set(AID)`, or `None` (representing a void, null, or missing value). For example, the pair `(get_next_trn, aid.loyalty1)` represents issuing the `get_next_trn` command with argument `loyalty1`; its response is represented by the pair `(get_next_trn_r, aids.`*`r`*`)`, where *`r`* is either `none` or `loyalty1`.

```
datatype CmdName = log_full
                 | get_next_trn | get_next_trn_r
                 | avail_funds  | avail_funds_r
```

```
datatype Value = aid.AID | aids.Set(AID) | none

nametype None    = { none }
nametype Command = ( CmdName, Value )
```

Continuing the previous example, the `spy` loyalty applet might want to pretend to be the `loyalty1` applet, in which case it would issue the spoofed message `loy_pur_int.spy.purse.(get_next_trn, loyalty1)`, with an expected response of `loy_pur_int.spy.purse.(get_next_trans_r, aid.`*`r`*`)`, where *r* is either `none` or `loyalty1`. From this example it is now clear how to construct the set of messages which are associated with each interface. First, the purse-to-loyalty interface has only one message that has no arguments or response, so it is represented by the singleton set containing the pair `(log_full, none)`.

```
nametype PurLoyMsg = { (log_full, none) }
```

Second, the loyalty-to-purse interface contains the commands representing the "get the next transaction" operation, namely `get_next_trn` and `get_next_trn_r`, which both take an applet identifier as a parameter. These commands are formalised by the following CSP, where the notation `{|aid|}` stands for all expansions of the `aid` value.

```
nametype LoyPurMsg = { (get_next_trn, l), (get_next_trn_r, l) | l <- {|aid|}
}
```

Last, the loyalty-to-loyalty interface contains the commands representing the available funds operation, namely `avail_funds` and `avail_funds_r`, which take an applet identifier and a set of applet identifiers as their parameters respectively. These are modelled in a similar manner to the loyalty-to-purse interface operations; the one complication is that the parameter types are different.

```
nametype LoyLoyMsg =
  { (avail_funds, l), (avail_funds_r, r) | l <- {|aid|}, r <- {|aids|} }
```

7.2.5.2 Inference Rules / Semantic Specification

Now that the main interfaces to the system have been defined, it is necessary to model their semantics by a set of inference rules. An inference rule is represented by a triple `(n, X, Y)`, where `n` represents the rule's unique name[11], `X` represents the set of capabilities required for the rule to fire, `Y` represents the set of capabilities gained from firing the rule. The inference rules, which represent the abstract semantics of the case study now follows, where each *operation* – a command with its response (if it has one) – is dealt with in turn.

First, the "log full" operation is considered. When the `log_full` command is issued, it enables the loyalty interfaces of each of the loyalties that have *registered* for the log full event; the idea being that the loyalties have a chance to check the transaction table, before entries that have not already been processed are overwritten. However, once a loyalty has been activated it may perform any of its operations. This is formally modelled by enabling the `get_next_trn` and `avail_funds` commands of that loyalty. Further, as the set of rules associated with the "log full"

[11] Uniqueness of rule names simplifies the techniques required to model some forms of fault injection.

operation is dependent on the set of *registered* loyalties, the formal model of the "log full" rules (`LogFullRules`) is parameterised by this set of loyalties (`RegLoys`). Hence, the set of inference rules that are associated with the log full operation is:

```
LogFullRules(RegLoys) =
{ ( log_full_op.l,
    { pur_loy_int.purse.l.(log_full, none) },
    { loy_pur_int.l.purse.(get_next_trans, aid.l3),
      loy_loy_int.l.l2.(avail_funds,       aid.l3) | l2 <- LoyaltyAID
                                                     l3 <- LoyaltyAID }
  ) |
  l  <- RegLoys
}
```

where the rule name `log_full_op.l`, is informally speaking[12] the name that is constructed by replacing `l` by a member of `RegLoys`.

Next, the rules associated with the "get next transaction" operation are discussed. Here the intention is that a loyalty can get the next *unprocessed* transaction from the purse applet's transaction list, by supplying the purse with its applet identifier. However, before scanning the transaction list, the authenticity of the request is checked by ensuring that the applet identifier belongs to the requesting applet. This is formally modelled by a special capability known as `self.l1.l2`, which enables an applet `l2` to "authentically" claim that it is applet `l1`.

```
channel self: AID . AID
```

Now the "get next transaction" rules can be formally defined by the `GetNextTrnRules` set, where each "get next transaction" command returns either the found transaction (`l2`) or the empty transaction (`aid.no_aid`). Further, the empty transaction is returned in two distinct circumstances: (1) for requests that fail authentication; and (2) for requests that have no *unprocessed* transaction within the purse's transaction list.

```
GetNextTrnRules =
let
  calc(l, l2) = ( if l == l2 then { l2, no_aid } else { no_aid } )
within
{ ( get_next_trn_op.l1.l2.l,
    { loy_pur_int.l1.purse.(get_next_trans,   aid.l2), self.l1.l },
    { loy_pur_int.l1.purse.(get_next_trans_r, aid.r) | r <- calc(l, l2) }
  ) |
  l1 <- LoyaltyAID,
  l2 <- LoyaltyAID,
  l  <- LoyaltyAID
}
```

Next, the rules for the "available funds" operation are discussed. Here the intention is to enable a loyalty to calculate the amount of funds that are available to it; that is its own funds, and the funds available to its co-operating loyalties. Thus, the rules associated with the "available funds" operation are dependent on the co-operation loyalty relation (`CoOpLoy`). These rules are formally modelled by the following CSP where: the operation is only enabled if the loyalty can

[12] It is better to delay the formal definition of the set of rule names until all the rules have been discussed.

authentically claim to be who it says it is; and the result is calculated by the `avial_funds_calc` function (which performs the appropriate recursive co-operating loyalty available funds lookup).[13]

```
AvailFundsRules(CoOpLoy) =
{ ( avail_funds_op.l1.l2.l,
    { loy_loy_int.l1.l2.(avail_funds,   aid.l), self.l1.l },
    { loy_loy_int.l1.l2.(avail_funds_r, aids.avial_funds_calc(l,l2,CoOpLoy))
}
  ) |
  l1 <- LoyaltyAID,
  l2 <- LoyaltyAID,
  l  <- LoyaltyAID
}
```

Having defined all the rules associated with each operation, it is now possible to provide a formal description of the set of rule names, where:

- the rules associated with the "log full" operation are parameterised by the registered loyalty identifier;
- the rules associated with the "get next" transaction operation are parameterised by the actual and claimed source loyalty identifiers, as well as another loyalty identifier used for authentication;
- the rules associated with the "available funds" operation are parameterised by the actual source and destination loyalty identifiers, as well as the claimed source loyalty identifier.

```
datatype RuleNames = log_full_op     . LoyaltyAID
                   | get_next_trn_op . LoyaltyAID . LoyaltyAID . LoyaltyAID
                   | avail_funds_op  . LoyaltyAID . LoyaltyAID . LoyaltyAID
```

It is now possible to formally define the set of inference rules that are associated with the case study, for any given set of registered loyalties and co-operating loyalties. This is simply the union of the rules associated with the case study operations.

```
RuleSet(RegLoys, CoOpLoy) =
  Union({ LogFullRules(RegLoys), GetNextTrnRules, AvailFundsRules(CoOpLoy) })
```

7.2.5.3 Analysing the Model

In order to perform the CSP vulnerability analysis of this model, the hazards (or critical information) need to be identified. In this case study there are two security-related types of critical information.

The first type of critical property is that a loyalty can never get a transaction that belongs to another loyalty; thus whenever a loyalty `l1` issues a get next transaction operation it must not

[13] The details of the "available funds calculation" operation are fairly straightforward but lengthy, due to the lack of built-in support for relations within the CSP/FDR language.

receive a transaction that contains a different loyalty's identifier (i.e., `l2 ≠ l1`). Formally this is modelled by the following CSP:

```
GetNextTrnCritInfo = { loy_pur_int.l1.purse.(get_next_trn_r, aid.l2)
                     | l1 <- LoyaltyAID,
                       l2 <- LoyaltyAID,
                       l2 != l1
                     }
```

The second type of critical property is that a loyalty should never be able to get funds from a loyalty that it does not co-operate with. In other words, the returned available funds to a loyalty `l`, as represented by a set of loyalty identifiers `L`, should not contain a loyalty identifier that does not belong to `l`'s set of co-operating loyalties `co_op_loy(l,CoOpLoy)`, where `CoOpLoy` represents the co-operating loyalty relationship. This is formally represented by the following CSP, where it is formally necessary to insist that a loyalty can co-operate with itself.

```
AvailFundsCritInfo(CoOpLoy) = { loy_loy_int.l.l1.(avail_funds_r, aids.L)
                              | l  <- LoyaltyAID,
                                l1 <- LoyaltyAID,
                                L  <- Set(LoyaltyAID),
                                diff(L,union({l},co_op_loy(l,CoOpLoy))) != {}
                              }
```

The information required for the vulnerability analysis is almost complete. All that remains is to:

1. choose an appropriate set of initial capabilities;
2. specify the set of registered loyalties (default all loyalties);
3. specify the co-operating loyalty relation;
4. link the resulting specification in with one of the general purpose CSP inference engines.

First, point 4 in the above list is addressed. The case study has been specified in a manner that enables it to be used with an existing general purpose CSP inference engine, which takes a collection of rules, some starting capabilities, and a set of *misuse* (hazardous) capabilities, and determines whether it is possible to reach a *misuse* capability. If it is possible to reach a misuse capability then the minimal sequence of inferences that led to that *misuse* can be displayed. Note that more advanced versions of the inference engine are able to cope with mode switching and interaction with external processes, such as for interacting with a state based model of the environment.

Next, points 1 and 2 are addressed. Recall that the "log full" operation enables all the operations of the registered loyalties; thus, if all loyalties are registered, then enabling the log full operation is sufficient to enable all of the operations within the case study. In addition to this, each loyalty applet should be able to authentically claim to be itself, so `self.l.l` is also enabled for all loyalty applets with applet identifier `l`.

Last, point 3 is addressed. This is slightly more complex as there is no obviously correct choice. For the purposes of discussion, two relations will be examined, namely the empty relation, where no loyalty co-operates with any other loyalty, and the relation `CoOpLoy_rel` where both the `spy` and `loyalty2` loyalties co-operate with `loyalty1`, but not each other.

```
CoOpLoy_rel = { (spy,      loyalty1), (loyalty1, spy),
                (loyalty1, loyalty2), (loyalty2, loyalty1)
              }
```

When the inference engine analysis is run with the above `CoOpLoy_rel` co-operating loyalty relationship, the inference engine identifies a misuse, which states that the `spy` can discover `loyalty2`'s available funds, even though it is not in its set of co-operating loyalties (i.e., `loyalty1`). This results from the transitive nature of the calculation of the available funds operation, where the funds available to `loyalty1` are also available to the `spy` (and `loyalty2`). Having identified the problem in the design, there are a number of solutions. For example, a simple restriction that the co-operating loyalties must be transitively closed is sufficient. Alternatively, the available funds calculation could be updated so that it only returns the funds of the loyalty that was explicitly asked. For the purposes, of the case study the first option is taken, as this requires no rewriting of the rules.

The second configuration of the inference engine, namely with the empty co-operating relation, is trivially transitively closed, and thus should not suffer from the transitive misuse problem identified in the previous paragraph. Indeed, when this configuration of the case study is analysed (model checked by FDR) no misuse can be found, and thus the design is not inherently "hazardous". Having determined that the design is not inherently "hazardous", how robust is it to faults in its implementation and use. The existing inference engines already have some ability to answer these questions, via two fault injection techniques: one which essentially adds n additional capabilities chosen from a faulty-capability set; and the other, which removes individual "firing" conditions from the rule. Following some experimentation with the fault sets, it quickly became established that providing the spy with the capability to authentically pretend to be another loyalty was critical; that is the spy could claim to be a different loyalty and pass the authentication check. For example if `self.spy.loyalty1`, is added to the initial set of capabilities, then it can quickly (by 2 rules) be inferred that the `spy` can get hold of a `loyalty1` transaction as represent by the `loy_pur_int.spy.purse(get_next_trn_r, aid.loyalty1)` capability. Further, it was established that the rules that check authentication were also vulnerable to the same type of failure, but purely for that rule.

Before leaving the case study, it should be noted that care must be taken in the construction of the rules, and fault models, in order to make the CSP/FDR vulnerability analysis tractable. Current, research activities are examining design patterns that aid in this task.

7.2.6 Summary

Once the correct specification has been constructed, using CSP, it can be incorporated into the B model. Using the techniques described elsewhere in this handbook, the B model is then refined down to code. Hence, where previously we were only able to use formal techniques in refining our specifications to code, we can now employ formal analysis in the initial construction of such specifications. In this way the use of CSP analysis to discover the critical properties of systems complements the B development method. We gain the benefits of formal analysis much earlier on in our code development cycle.

It isn't obvious how one would use a theorem prover to achieve such results. Undoubtedly, it must be possible to devise an applicable method. However, model-checking is a theory which is mature and well suited to this type of analysis. The process of specification refinement discussed

here fits easily into the body of knowledge already present within the community. Our effort is focused on constructing the models, as it should be, and not on finding a way to make the B formalism meet our needs.

8 Summary

This handbook is intended to provide work practices and guidelines to software practitioners in using a formal development methodology. This methodology is centred around the B method, and it provides guidelines on how to use B (or a combination of B with other notations) in the various stages in the development lifecycle:

- Modelling notations: Besides presenting the B method, this section describes: event B that extends B to take into account abstract systems, and; the UML+B tool that converts UML class diagrams and state charts into B.

- Constructing models: In this section we present two approaches on how to model a system: in the system-level model one constructs a formal model of the overall system; while in a computational model one constructs only a formal model of the part of the system we want study, and then verify that it interacts correctly with the overall system.

- Analysing models: This section shows different ways in which formal models may be analysed for validity and soundness. We start by presenting a procedure for the review of B models by domain experts. Next, we show how B can be used to verify safety and identify vulnerabilities. Last, we describe strategies to prove (within Atelier-B) *proof obligations*, and in this way validate a model.

- Refining models: This section describes how an abstract model can be successively refined until a correct system implementation is obtained. For complex models, in addition to refinement is necessary to use decomposition - where the overall system is partitioned into several sub-systems. Furthermore, this section gives guidance on how to prove refinement and how to refine UML+B diagrams.

- Implementing models: This section provides guidelines on automatically generating code from B implementations. Moreover, it discusses the generation of code that must comply with timing and resources constraints. Another important issue discussed is the integration of formal code with existing non-formal code.

The MATISSE project aims to exploit and enhance existing generic methodologies and associated technologies that support the correct construction of software-based systems. In particular, a strong emphasis was placed on the use of the B Method, and on integrating B with other methods, namely UML and CSP. The techniques developed by MATISSE were driven by three major industrial case studies: transportation, healthcare, and smart card.

A The specification of the Analyser

A.1 Machine ANALYSER.mch

We give the abstract specification of the Analyser in ANALYSER.mch. This machine is derived from the basic UML statechart-diagram.

The variables of the abstract system are astate and acmd, modelling the state of the Analyser and the command given to the Analyser, respectively.

```
MACHINE ANALYSER
SEES
   DEF0
INCLUDES
   GLOBALVAR_PLATE,
   R_PROC,
   A_PROC
SETS
   ASTATE={aidle,amove,aanalyse,arecieve,adeliver,aabort,asuspended}
VARIABLES
   astate,
   acmd
INVARIANT
   astate ∈ ASTATE ∧
   acmd ∈ ACOMMAND
INITIALISATION
   astate := aidle ||
   acmd := receive

OPERATIONS
NewCommand (cc) =
   PRE
      cc ∈ ACOMMAND
   THEN
      SELECT astate=aidle THEN acmd:=cc END
   END;

MoveA (pos) =
   PRE
      pos ∈ NAT
   THEN
      SELECT astate=aidle ∧
             acmd=move
      THEN   astate:=amove ||
             Safe
      END
   END;

MoveAFail  =
   BEGIN
      SELECT astate=aidle ∧
             acmd=move
      THEN   astate:=asuspended ||
             NotSafe
      END
   END;
```

```
MoveAMinRemedy  =
   BEGIN
      SELECT astate=asuspended ∧
             acmd=move
      THEN   astate:=aidle ||
             Safe
      END
   END;

MoveAMidRemedy  =
   BEGIN
      SELECT astate=asuspended ∧
             acmd=move
      THEN   astate:=aidle ||
             Safe
      END
   END;

MoveAReady (pos) =
   PRE
      pos ∈ NAT
   THEN
      SELECT astate=amove ∧
             acmd=move
      THEN   astate:=aidle
      END
   END;

MoveANotReady (pos) =
   PRE
      pos ∈ NAT
   THEN
      SELECT astate=amove ∧
             acmd=move
      THEN   astate:=asuspended
      END
   END;

MoveRemedy (pos) =
   PRE
      pos ∈ NAT
   THEN
      SELECT astate=asuspended ∧
             acmd=move
      THEN   astate:=amove ||
             Safe
      END
   END;

Analyse  =
   BEGIN
      SELECT 0=0
      THEN   astate:=aanalyse ||
             PlateInAnalyser ||
         SELECT astate=aidle ∧ acmd=analyse THEN skip END
      END
   END;

AnalyseFail  =
   BEGIN
```

```
        SELECT astate=aidle ∧
               acmd=analyse
        THEN   astate:=asuspended
        END
    END;

AnalysePlateFail  =
    BEGIN
        SELECT 0=0
        THEN   astate:=asuspended ||
               NoPlateInAnalyser ||
           SELECT astate=aidle ∧ acmd=analyse THEN skip END
        END
    END;

AnalysePosRemedy  =
    BEGIN
        SELECT astate=asuspended ∧
               acmd=analyse
        THEN   astate:=aidle ||
               Safe
        END
    END;

AnalyseMaxRemedy  =
    BEGIN
        SELECT astate=asuspended ∧
               acmd=analyse
        THEN   astate:=aidle
        END
    END;

AnalyseReady  =
    BEGIN
        SELECT astate=aanalyse ∧
               acmd=analyse
        THEN   astate:=aidle
        END
    END;

AnalyseNotReady  =
    BEGIN
        SELECT astate=aanalyse ∧
               acmd=analyse
        THEN   astate:=asuspended
        END
    END
;
AnalyseRemedy  =
    BEGIN
        SELECT astate=asuspended ∧
               acmd=analyse
        THEN   astate:=aanalyse ||
               PlateInAnalyser
        END
    END;

Deliver  =
    BEGIN
        SELECT 0=0
```

```
      THEN   astate:=adeliver ||
             PlateInAnalyser || ReadyFor_blank ||
             SELECT astate=aidle ∧ acmd=deliver THEN skip END
      END
   END;

DeliverFail  =
   BEGIN
      SELECT astate=aidle ∧
             acmd=deliver
      THEN   astate:=asuspended
      END
   END;

DeliverPlateFail  =
   BEGIN
      SELECT 0=0
      THEN   astate:=asuspended ||
             NoPlateInAnalyser ||
             SELECT astate=aidle ∧ acmd=deliver THEN skip END
      END
   END;

DeliverFailAtRobot  =
   BEGIN
      SELECT astate=aidle ∧
             acmd=deliver
      THEN   astate:=asuspended ||
             NotReadyFor_blank
      END
   END;

DeliverPosRemedy  =
   BEGIN
      SELECT astate=asuspended ∧
             acmd=deliver
      THEN   astate:=aidle ||
             Safe
      END
   END;

DeliverReady  =
   BEGIN
      SELECT astate=adeliver ∧
             acmd=deliver
      THEN   astate:=aidle ||
             NoPlateInAnalyser
      END
   END;

DeliverNotReady  =
   BEGIN
      SELECT astate=adeliver ∧
             acmd=deliver
      THEN   astate:=asuspended
      END
   END;

DeliverNotReadyPlate  =
   BEGIN
```

```
        SELECT astate=adeliver ∧
               acmd=deliver
        THEN   astate:=asuspended ||
               PlateInAnalyser
        END
    END;

DeliverRemedy  =
    BEGIN
        SELECT astate=asuspended ∧
               acmd=deliver
        THEN   astate:=adeliver ||
               Safe || NoPlateInAnalyser
        END
    END;

Receive  =
    BEGIN
        SELECT 0=0
        THEN   astate:=arecieve ||
               NoPlateInAnalyser || Available_blank ||
               SELECT astate=aidle ∧ acmd=receive THEN skip END
        END
    END;

ReceiveFail  =
    BEGIN
        SELECT astate=aidle ∧
               acmd=receive
        THEN   astate:=asuspended
        END
    END;

ReceivePlateFail  =
    BEGIN
        SELECT 0=0
        THEN   astate:=asuspended ||
               PlateInAnalyser ||
           SELECT astate=aidle ∧ acmd=receive THEN skip END
        END
    END;

ReceiveFailAtRobot  =
    BEGIN
        SELECT astate=aidle ∧
               acmd=receive
        THEN   astate:=asuspended ||
               NoAvailable_blank
        END
    END;

ReceivePosRemedy  =
    BEGIN
        SELECT astate=asuspended ∧
               acmd=receive
        THEN   astate:=aidle ||
               Safe
        END
    END;
```

```
ReceiveReady  =
   BEGIN
      SELECT astate=arecieve ∧
             acmd=receive
      THEN   astate:=aidle ||
             PlateInAnalyser
      END
   END;

ReceiveNotReady  =
   BEGIN
      SELECT astate=arecieve ∧
             acmd=receive
      THEN   astate:=asuspended
      END
   END;

ReceiveNotReadyPlate  =
   BEGIN
      SELECT astate=arecieve ∧
             acmd=receive
      THEN   astate:=asuspended ||
             NoPlateInAnalyser
      END
   END;

ReceiveRemedy  =
   BEGIN
      SELECT astate=asuspended ∧
             acmd=receive
      THEN   astate:=arecieve ||
             Safe || PlateInAnalyser
      END
   END;

AServiceNotReady  =
   BEGIN
      SELECT astate=aidle
      THEN   astate:=asuspended
      END
   END;

ARemedy  =
   BEGIN
      SELECT astate=asuspended
      THEN   astate:=aidle
      END
   END;

AFailure  =
   BEGIN
      SELECT astate=asuspended
      THEN   astate:=aabort
      END
   END
END
```

A.2 Machine `DEF0.mch`

This machine defines the commands used in the Analyser.

```
MACHINE DEF0
SETS
   ACOMMAND={move,analyse,receive,deliver}
END
```

A.3 Machine GLOBALVAR_PLATE.mch

The global variable plate is given in a separate machine GLOBALVAR_PLATE.mch.

```
MACHINE GLOBALVAR_PLATE
VARIABLES
   plate
INVARIANT
   plate ∈ BOOL
INITIALISATION
   plate := FALSE

OPERATIONS
ChangePlate (bb) =
   PRE
      bb ∈ BOOL
   THEN
      plate:=bb
   END;

PlateInAnalyser  =
   BEGIN
      SELECT plate=TRUE THEN skip END
   END;

NoPlateInAnalyser  =
   BEGIN
      SELECT plate=FALSE THEN skip END
   END
END
```

A.4 Machine A_PROC.mch

The exported global of the Analyser are given in A_PROC.mch.

```
MACHINE A_PROC

OPERATIONS
ReadyToDeliver  =
   BEGIN skip END;

NotReadyToDeliver  =
   BEGIN skip END;

ReadyToReceive  =
   BEGIN skip END;

NotReadyToReceive  =
   BEGIN skip END;

BlankInAnalyser  =
   BEGIN skip END;

NoBlankInAnalyser  =
   BEGIN skip END
END
```

A.5 Machine R_PROC.mch

The imported procedures from the Robot are given in R_PROC.mch. These procedures are only given as skip, since the Analyser need not be aware of their implementation.

```
MACHINE R_PROC

OPERATIONS
Safe  =
   BEGIN skip END;

NotSafe  =
   BEGIN skip END;

Available_blank  =
   BEGIN skip END;

NotAvailable_blank  =
   BEGIN skip END;

ReadyFor_blank  =
   BEGIN skip END;

NotReadyFor_blank  =
   BEGIN skip END
END
```

B The Language CSP_M

CSP_M is the machine readable dialect of CSP [Roscoe98,FDR97], used to define both the specification and implementation of models verified by FDR. It is very close to pure CSP and it is embedded in a Gofer/Haskell like functional programming language, which augments its utility substantially. A brief description of the key features of the CSP_M syntax follows, where the formal syntax is in a `typewriter` font and keywords and symbols are highlighted in **bold**.

`channel` `channel_id : ChannelType` is a CSP_M declaration of a **`channel`**, which is idenitified by its `channel_id` and communicates values of type `ChannelType`.

`datatype` `name = string1 . X1`$_1$ `. ... . X1`$_n$ `| ... | stringN . XN`$_1$ `. ... . XN`$_m$, where `name` and `string1` to `stringN` represent globally unique identifiers, and `XI`$_J$ represents either explicit or named sets of data. Note that CSP_M has two inbuilt datatypes, namely `Bool` and `Int`, representing booleans (true, false) and integers between $\pm(2^{31}-1)$.

`nametype` `name = X`, where a set expression (`X`) is given a `name` which is intended to be used as a type.

`STOP` is the simplest CSP process; it never engages in any action, and never (successfully) terminates; that is it cannot be used as a successful termination for sequential composition.

`SKIP` is the process that terminates successfully, and is the unit of sequential composition (i.e. `P ; SKIP = P = SKIP ; P`). Note that the precise formal semantics of **`SKIP`** is surprisingly complex.

`a -> P` is the most basic program constructor. It waits to perform the event a and after this has occurred subsequently behaves as process `P`. The same notation is used for outputs `(c!v -> P)`, inputs `(c?x -> P(x))` and simple communication `(c.x -> P(x))` of values along named channels.

`bool_exp & a -> P` is the process which behaves as `a -> P` if the `bool_exp` is true.

`P |~| Q` represents the non-deterministic or internal choice between `P` and `Q`.

`P [] Q` represents the external choice between `P` and `Q`, which is deterministic so long as the set of initial events of P and Q are disjoint. The process `P [] Q` offers the initial actions of both `P` and `Q` to its environment; its subsequent behaviour is like `P` if the initial action chosen was possible only for `P` and like `Q` if the action selected `Q`. If both `P` and `Q` have common initial actions, its subsequent behaviour is non-deterministic (like `|~|`). Note that **`STOP`** is the unit of external choice (i.e. `P [] STOP = P = STOP [] P`).

`[] x : X @ x -> P` represents replicated external choice.

`P \ A` is the CSP hiding operator. This process behaves as `P` except that events in set A are hidden from the environment and are solely (and internally) determined by `P`; the

environment can neither observe nor influence them. `P \ {|channel_id|}` hides all communications along the channel `channel_id`.

`P [| A |] Q` represents shared parallel (concurrent) composition. `P` and `Q` evolve separately, except that events from A occur only when `P` and `Q` agree (i.e. synchronise) to perform them.

`P [Alphabet1 || Alphabet2] Q` represents alphabetised parallel composition. Any events in `P` which are not in `Alphabet1` are blocked from occurring and similarly for `Q`. `P` and `Q` evolve separately except or events in the intersection of `Alphabet1` and `Alphabet2` occur only when `P` and `Q` agree (i.e. synchronise) to perform them.

`|| i : Indexing_set @ [A(i)] P(i)` represents alphabetised parallel composition. For example `|| i : {1,2,3} @ [A(i)] P(i)` is equivalent to `(P(1) [A(1) || A(2)] P(2)) [Union(A(1),A(2))|| A(3)] P(3).`

`P [[ channel_id1 <-> channel_id2 ]] Q` represents linked parallel composition. The two processes communicate via a common new hidden channel resulting from a renaming of `channel_id1` and `channel_id2`.

`{| name |}` represnets the set of all expansions of `name`; that is the set containing `{ name.x }` for all values `x` that can be prefixed by "`name.`".

`diff(X, Y)` represnets the set difference operation.

`union(X, Y)` represnets the set union operation.

`Union({X1, …, Xn})` represnets the multi argument (distruibuted) set union operation; that is the union of sets `X1, …, Xn`.

`Set(X)` represnets the power set operation; that is the set of all subsets of `X`.

`if bool_exp then exp1 else exp2` represents the conditional evaluation of an expresion, where `exp1` is returned if the `bool_exp` is true, and `exp2` if `bool_exp` is false.

`let name=val within body` represents a technique for local aliasing, where each occurrence of `name` within the `body` is replaced with `val`. Note that the name can optionally include have a bracketted set of comma separated arguments, in which case a local function has been declared.

`chase(P)` The process `P` with the 'chase' operator applied to it. As a consequence, internal events, called tau-events, in `P`, are give priority over all other events. Where there is an option to perform several tau-events, one of them is chosen arbitrarily. Note the use of chase changes the semantics of the process `P`, if the order in which the hidden tau-events take place matters.

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