A Tutorial on (Co)Algebras and (Co)Induction*

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Abstract. *Algebraic structures which are generated by a collection of constructors—like natural numbers (generated by a zero and a successor) or finite lists and trees—are of well-established importance in computer science. Formally, they are initial algebras. Induction is used both as a definition principle, and as a proof principle for such structures. But there are also important dual "coalgebraic" structures, which do not come equipped with constructor operations but with what are sometimes called "destructor" operations (also called observers, accessors, transition maps, or mutators). Spaces of infinite data (including, for example, infinite lists, and non-well-founded sets) are generally of this kind. In general, dynamical systems with a hidden, black-box state space, to which a user only has limited access via specified (observer or mutator) operations, are coalgebras of various kinds. Such coalgebraic systems are common in computer science. And "coinduction" is the appropriate technique in this coalgebraic context, again both as a definition principle and as a proof principle. The latter involves bisimulations. It is the aim of this tutorial to provide a brief introduction to this relatively new field of coalgebra.*

1 Introduction

Algebra is a well-established part of mathematics, dealing with sets with operations satisfying certain properties, like groups, rings , vector spaces, etcetera. Its results are essential throughout mathematics and other sciences. *Universal* algebra is a part of algebra in which algebraic structures are studied at a high level of abstraction and in which general notions like homomorphism, subalgebra, congruence are studied in themselves, see e.g. [17, 48, 70]. A further step up the abstraction ladder is taken when one studies algebra with the notions and tools from category theory. This approach leads to a particularly concise notion of what is an algebra (for a functor or for a monad), see for example [45]. The conceptual world that we are about to enter owes much to this categorical view, but it also takes inspiration from universal algebra, see e.g. [65].

In general terms, a program in some programming language manipulates data. During the development of computer science over the past few decades it became clear that an abstract description of these data is desirable, for example to ensure that one's program does not depend on the particular representation of the data on which it operates. Also, such abstractness facilitates correctness proofs. This desire led to the use of algebraic methods in computer science, in a branch called *algebraic specification* or *abstract data type theory*. The object of study are data types in themselves, using notions of techniques which are familiar from algebra. The data types used by computer scientists are often generated from a given collection of (constructor) operations,

*This paper is published in: EATCS Bulletin 62 (1997), p.222–259.

and it is for this reason that "initiality" of algebras plays such an important role (as first clearly emphasised in [26]). See for example [19, 71, 70] for more information.

Standard algebraic techniques have proved useful in capturing various essential aspects of data structures used in computer science. But it turned out to be difficult to algebraically describe some of the inherently dynamical structures occuring in computing. Such structures usually involve a notion of state, which can be transformed in various ways. Formal approaches to such state-based dynamical systems generally make use of automata or transition systems, see e.g. [60, 53, 50] as classical early references. During the last decade the insight gradually grew that such state-based systems should not be described as algebras, but as so-called coalgebras. These are the formal duals of algebras, in a way which will be made precise in this tutorial. The dual property of initiality for algebras, namely finality turned out to be crucial for such coalgebras. And the logical reasoning principle that is needed for such final coalgebras is not induction but coinduction. There is no single reference in which this link between state-based dynamical systems and coalgebra is made explicitly (with all its ramifications), but important insights can be found in [21, 69, 5, 25, 40, 1, 2, 51, 15, 58, 63, 64, 61, 20, 37, 35, 34, 67, 65, 38]. This list is incomplete and does not do justice to the various contributors to this area, but it hopefully gives the reader an impression of some of the developments.

These notions of coalgebra and coinduction are still relatively unfamiliar, and it is our aim in this tutorial to explain them in elementary terms. There is currently little introductory material available, since most of the literature already assumes some form of familiarity either with category theory, or with the (dual) coalgebraic way of thinking (or both). The author's experiences in lecturing about coalgebras is that the material in itself is usually not seen as difficult, but that it takes a subtle change-of-view (with respect to the traditional algebraic approach) to be able to appreciate, recognise and apply the coalgebraic notions and techniques.

Before we start, we should emphasise that there is no new (research) material in this tutorial. Everything that we present is either known in the literature, or in the folklore, so we do not have any claims to originality (except possibly regarding the presentation of the material). Also, our main concern is with conveying ideas, and not with giving a correct representation of the historical developments of these ideas. Reference are given mainly in order to provide sources for more (background) information.

Also, we should emphasise that we do not assume any knowledge of category theory on the part of the reader. We shall often use the diagrammatic notation which is typical of category theory, but only in order to express equality of two composites of functions, as often used also in other contexts. This is simply the most efficient and most informative way of presenting such information. But in order to fully appreciate the underlying duality between algebra and induction on the one hand, and coalgebra and coinduction on the other, some elementary notions from category theory are needed, especially the notions of *functor* (homomorphism of categories), and of *initial* and *final* (also called *terminal*) object in a category. Here we shall explain these notions in the concrete set-theoretic setting in which we are working, but we definitely encourage the interested reader who wishes to further pursue the topic of this tutorial to study category theory in greater detail. Among the many available texts on category theory, [57, 68, 4] are recommended as easy-going starting points, [9, 18, 41] as more substantial texts, and [42, 13] as advanced reference texts.

This tutorial starts with some introductory expositions in Sections 2 – 4. The technical material in the subsequent sections is organised as follows.

1. The starting point is ordinary induction, both as a definition principle and as a proof principle. We shall assume that the reader is familiar with induction, over natural numbers, but also over other data types, say of lists, trees or (in general) of terms. The first real step is to reformulate ordinary induction in a more abstract way, using initiality (see Section 5). More precisely, using initiality for "algebras of a functor". This is something which we do not assume to be familiar. We therefore explain how signatures of operations give rise to certain functors, and how algebras of these functors correspond to algebras (or models) of the signatures (consisting of a set equipped with certain functions interpreting the operations). This description of induction in terms of algebras (of functors) has the advantage that it is

highly uniform, in the sense that it applies in the same way to all kinds of (algebraic) data types. Further, it can be dualised easily, thus giving rise to the theory of coalgebras.

2. The dual notion of an algebra (of a functor) is a coalgebra (of a functor). It can also be understood as a model consisting of a set with certain operations, but the direction of these operations is not as in algebra. The dual notion of initiality is finality, and this finality gives us coinduction, both as a definition principle and as a reasoning principle. This pattern is as in the previous point, and is explained in Section 6.

3. Finally in Section 7, we give an alternative formulation of the coinductive reasoning principle (introduced in terms of finality) which makes use of bisimulations. These are relations on coalgebras which are suitably closed under the (coalgebraic) operations; they may be understood as duals of congruences, which are relations which are closed under algebraic operations. Bisimulation arguments are used to prove the equality of two elements of a final coalgebra, and require that these elements are in a bisimulation relation. Such arguments are much used in concurrency theory. We conclude with a brief discussion of the various predicates and relations which are of relevance in algebra and coalgebra.

In a first approximation, the duality between induction and coinduction that we intend to describe can be understood as the duality between least and greatest fixed points (of a monotone function). These notions generalise to least and greatest fixed points of a functor, which are suitably described as initial algebras and final coalgebras. The point of view mentioned in 1. and 2. above can be made more explicit as follows—without going into technicalities yet. The abstract reformulation of induction that we will describe is:

$$\boxed{\text{induction} \;=\; \text{use of initiality for algebras}}$$

An algebra (of a certain kind) is *initial* if for an arbitrary algebra (of the same kind) there is a unique homomorphism (structure-preserving mapping) of algebras:

$$\left(\begin{array}{c}\text{initial}\\ \text{algebra}\end{array}\right) \dashrightarrow^{\text{unique}\atop\text{homomorphism}} \left(\begin{array}{c}\text{arbitrary}\\ \text{algebra}\end{array}\right) \tag{1}$$

This principle is extremely useful. Once we know that a certain algebra is initial, by this principle we can define functions acting on this algebra. Initiality involves unique existence, which has two aspects:

Existence. This corresponds to (ordinary) *definition* by induction.

Uniqueness. This corresponds to *proof* by induction. In such uniqueness proofs, one shows that two functions acting on an initial algebra are the same by showing that they are both homomorphisms (to the same algebra).

The details of this abstract reformulation will be elaborated as we proceed.

Dually, coinduction may be described as:

$$\boxed{\text{coinduction} \;=\; \text{use of finality for coalgebras}}$$

A coalgebra (of some kind) is *final* (or *terminal*) if for an arbitrary coalgebra (of the same kind), there is a unique homomorphism of coalgebras as shown:

$$\left(\begin{array}{c}\text{arbitrary}\\ \text{coalgebra}\end{array}\right) \dashrightarrow^{\text{unique}\atop\text{homomorphism}} \left(\begin{array}{c}\text{final}\\ \text{coalgebra}\end{array}\right) \tag{2}$$

Again we have the same two aspects: existence and uniqueness, corresponding this time to definition and proof by coinduction.

The initial algebras and terminal coalgebras which play such a prominent role in this theory can be described in a canonical way: an initial algebra can be obtained from the closed terms (i.e. from those terms which are generated by iteratively applying the algebra's constructor operations), and the terminal coalgebra can be obtained from the pure observations. The latter is probably not very familiar, and will be illustrated in several examples in the next section.

2 Algebraic and coalgebraic phenomena

The distinction between algebra and coalgebra pervades computer science and has been recognised by many people in many situations, usually in terms of data versus machines. A modern, mathematically precise way to express the difference is in terms of algebras and coalgebras. The basic dichotomy may be described as *construction* versus *observation*. It may be found in process theory [50], data type theory [21, 25, 5, 40] (including the theory of classes and objects in object-oriented programming [61, 30, 37, 35]), semantics of programming languages [46] (denotational versus operational [64, 67, 6]) and of lambda-calculi [58, 59, 20, 31], automata theory [53], system theory [65, 34], natural language theory [10, 62] and many other fields.

We assume that the reader is familiar with definitions and proofs by (ordinary) induction. As a typical example, consider for a fixed data set A, the set $A^\star = \mathsf{list}(A)$ of finite sequences (lists) of elements of A. One can inductively define a length function $\mathsf{len}\colon A^\star \to \mathbb{N}$ by the two clauses:

$$\mathsf{len}(\langle\rangle) = 0 \qquad \text{and} \qquad \mathsf{len}(a \cdot \sigma) = 1 + \mathsf{len}(\sigma)$$

for all $a \in A$ and $\sigma \in A^\star$. Here we have used the notation $\langle\rangle \in A^\star$ for the empty list (sometimes called nil), and $a \cdot \sigma$ (sometimes written as $\mathsf{cons}(a, \sigma)$) for the list obtained from $\sigma \in A^\star$ by prefixing $a \in A$. As we shall see later, the definition of this length function $\mathsf{len}\colon A^\star \to \mathbb{N}$ can be seen as an instance of the above initiality diagram (1).

A typical induction proof that a predicate $P \subseteq A^\star$ holds for all lists requires us to prove the induction assumptions

$$P(\langle\rangle) \qquad \text{and} \qquad P(\sigma) \Rightarrow P(a \cdot \sigma)$$

for all $a \in A$ and $\sigma \in A^\star$. For example, in this way one can prove that $\mathsf{len}(\sigma \cdot a) = 1 + \mathsf{len}(\sigma)$ by taking $P = \{\sigma \in A^\star \mid \forall a \in A.\, \mathsf{len}(\sigma \cdot a) = 1 + \mathsf{len}(\sigma)\}$. (Essentially, this induction proof method says that $A^\star$ has no proper subalgebras.) In this (algebraic) setting we make essential use of the fact that all finite lists of elements of A can be constructed from the two operations $\mathsf{nil} \in A^\star$ and $\mathsf{cons}\colon A \times A^\star \to A^\star$. As above, we also write $\langle\rangle$ for nil and $a \cdot \sigma$ for $\mathsf{cons}(a, \sigma)$.

Next we describe some typically coalgebraic phenomena, by sketching some relevant examples. Many of the issues that come up during the description of these examples will be explained in further detail in later sections.

(i) Consider a black-box machine (or process) with one (external) button and one light. The machine performs a certain action only if the button is pressed. And the light goes on only if the machine stops operating (i.e. has reached a final state); in that case, pressing the button has no effect any more. A client on the outside of such a machine cannot directly observe the internal state of the machine, but (s)he can only observe its behaviour via the button and the light. In this simple (but paradigmatic) situation, all that can be observed directly about a particular state of the machine is whether the light is on or not. But a user may iterate this experiment, and record the observations after a change of state caused by pressing the button[1]. In this situation, a user can observe how many times (s)he has to press the button to make the light go on. This may be zero times (if the light is already on), $n \in \mathbb{N}$ times, or infinitely many times (if the machine keeps on operating and the light never goes on).

Mathematically, we can describe such a machine in terms of a set X, which we understand as the unknown state space of the machine, on which we have a function

$$\mathsf{button}\colon X \longrightarrow \{*\} \cup X$$

where $*$ is a new symbol not occurring in X. In a particular state $s \in X$, applying the function button—which corresponds to pressing the button—has two possible outcomes: either $\mathsf{button}(s) = *$, meaning that the machine stops operating and that the light goes on, or $\mathsf{button}(s) \in X$. In the latter case the machine has moved to a next state as a result of the button being pressed. (And

[1] It is assumed that such actions of pressing a button happen instantaneously, so that there is always an order in the occurrence of such actions.

in this next state, the button can be pressed again) The above pair $(X, \mathsf{button}\colon X \to \{*\} \cup X)$ is an example of a coalgebra.

The observable behaviour resulting from iterated observations as described above yields an element of the set $\overline{\mathbb{N}} = \mathbb{N} \cup \{\infty\}$, describing the number of times the button has to be pressed to make the light go on. Actually, we can describe this behaviour as a function $\mathsf{beh}\colon X \to \overline{\mathbb{N}}$. As we shall see later, it can be obtained as instance of the finality diagram (2).

(ii) Let us consider a slightly different machine with two buttons: value and next. Pressing the value button results in some visible indication (or attribute) of the internal state (e.g. on a display), taking values in a dataset A, without affecting the internal state. (Hence pressing value twice consecutively yields the same result.) By pressing the next button the machine moves to another state (the value of which can be inspected again). Abstractly, this new machine can be described as a coalgebra

$$\langle \mathsf{value}, \mathsf{next} \rangle\colon X \longrightarrow A \times X$$

on a state space X. The behaviour that we can observe of such machines is the following: read the value after pressing the next button $n \in \mathbb{N}$ times. This results in an infinite sequence $(a_0, a_1, a_2, \ldots) \in A^{\mathbb{N}}$ of elements of the dataset A, with element a_i describing the value after pressing next i times. Observing this behaviour for every state $s \in X$ gives us a function $\mathsf{beh}\colon X \to A^{\mathbb{N}}$, which can be described as instance of (2).

(iii) The previous example is leading us in the direction of a coalgebraic description of classes in object-oriented languages. Suppose we wish to capture the essential aspects of the class of points in a (real) plane that can be moved around by a client. In this situation we certainly want two attribute buttons $\mathsf{first}\colon X \to \mathbb{R}$ and $\mathsf{second}\colon X \to \mathbb{R}$ which tell us, when pushed, the first and second coordinate of a point belonging to this class. As before, the X plays the role of a hidden state space, and elements of X are seen as objects of the class (so that an object is identified with a state). Further we want a button (or method, in object-oriented terminology) $\mathsf{move}\colon X \times (\mathbb{R} \times \mathbb{R}) \to X$ which requires two parameters (corresponding to the change in first and second coordinate). This move operation allows us to change a state in a certain way, depending on the values of the parameters. The move method can equivalently be described as a function $\mathsf{move}\colon X \to X^{(\mathbb{R} \times \mathbb{R})}$ taking the state as single argument, and yielding a function $(\mathbb{R} \times \mathbb{R}) \to X$ from parameters to states.

As a client of such a class we are not interested in the actual details of the implementation (what the state space X exactly looks like) as long as the behaviour is determined by the following two equations:

$$\begin{aligned} \mathsf{first}(\mathsf{move}(s, (d1, d2))) &= \mathsf{first}(s) + d1 \\ \mathsf{second}(\mathsf{move}(s, (d1, d2))) &= \mathsf{second}(s) + d2 \end{aligned}$$

These describe the first and second coordinates after a move in terms of the original coordinates and the parameters of the move. Such equations can be seen as constraints on the observable behaviour.

An important aspect of the object-oriented approach is that classes are built around a hidden state space, which can only be observed and modified via certain specified operations. A user is not interested in the details of the actual implementation, but only in the behaviour that is realised. This is why our black-box description of classes with an unknown state space X is appropriate.

The three buttons of such a class (as abstract machine) can be combined into a single function

$$\langle \mathsf{first}, \mathsf{second}, \mathsf{move} \rangle\colon X \longrightarrow \mathbb{R} \times \mathbb{R} \times X^{(\mathbb{R} \times \mathbb{R})}$$

which forms a coalgebra on the state space X. The observable behaviour is very simple in this case. It consists of the values of the first and second coordinates, since if we know these values, then we know the future observable behaviour: the only change of state that we can bring about is through the move button; but its observable effect is determined by the above two equations. Thus what we can observe about a state is obtained by direct observation, and repeated observations do not produce new information. Hence our behaviour function takes the form $\mathsf{beh}\colon X \to \mathbb{R} \times \mathbb{R}$,

and is again an instance of (2). In automata-theoretic terms one can call the space $\mathbb{R} \times \mathbb{R}$ the minimal realisation (or implementation) of the specified behaviour.

(iv) We return to the second example with two buttons $(\mathsf{value}, \mathsf{next})\colon X \to A \times X$. This example may also be understood as a deterministic transition system: write for states $s, s' \in X$ and for an observable value $a \in A$,

$$s \xrightarrow{a} s' \qquad \text{if and only if} \qquad \mathsf{value}(s) = a \text{ and } \mathsf{next}(s) = s'.$$

We read $s \xrightarrow{a} s'$ as: in state s we can observe a and move on to s'. The trace $\mathrm{Tr}(s)$ of observations of a state $s \in X$ in this transition system is

$$\mathrm{Tr}(s) = (a_1, a_2, \ldots) \qquad \text{where} \qquad s \xrightarrow{a_1} s_1 \xrightarrow{a_2} s_2 \cdots$$

It is the sequence of the observable behaviour $\mathsf{beh}(s) \in A^{\mathbb{N}}$ of the state $s \in X$, as identified in (ii). The successor states s_i of s are completely determined because we have a deterministic machine in which we allow only one next state.

(v) Consider a general, not necessarily deterministic, transition system $(X, A, \longrightarrow)$, where $\longrightarrow$ is a subset of $X \times A \times X$. It can equivalently be described in coalgebra form using a powerset $\mathcal{P}$, namely, as a function

$$\alpha\colon X \longrightarrow \mathcal{P}(A \times X) \cong \mathcal{P}(X)^A$$

The equivalence is based on: $(a, t) \in \alpha(s) \Leftrightarrow s \xrightarrow{a} t$. The (deterministic) transition system mentioned in the previous point is a special case with each "successor set" $\alpha(s)$ a singleton. Such a general transition system α may also be understood as a (transition function of a) non-deterministic automaton with the set A as input alphabet.

The labels $a \in A$ of such a transition system are seen as observable. What then, is an appropriate domain of observations? This is a difficult question. We shall not give a complete answer, but mention some of the difficulties, sketch solutions, and give references for further details. First, we consider a variation of the above kind of transition systems, given by maps of the form:

$$\alpha\colon X \longrightarrow (A \times X)^\star = \mathsf{list}(A \times X)$$

In this case we can move from a state $s \in X$ to an ordered sequence $\alpha(s) \in (A \times X)^\star$ of labels and next states, say of the form $\alpha(s) = \langle (a_1, s_1), \ldots, (a_n, s_n) \rangle$. And we can continue this process with all these s_i's. The resulting space of observations consists of all possibly infinite A-labeled trees with at each node finitely many ordered branches. We shall return to such coalgebras in Example 6.7.

In a next step, consider transition systems of the form

$$\alpha\colon X \longrightarrow \mathcal{P}_f(A \times X)$$

which are finitely non-deterministic, in the sense that for each set $s \in X$ there are only finitely many pairs (a, t) with $s \xrightarrow{a} t$ (which are no longer ordered). One might think that the space of observations is now simply the set of all (possibly infinite) A-labeled trees with finitely many (non-ordered) branches, but the situation is more subtle. The problem is that for a state s there may be two transitions $s \xrightarrow{a} t_1$ and $s \xrightarrow{a} t_2$ with the same label, for which it is not clear yet if they lead to the same observations (and hence should be identified). In the previous situation (with $(A \times X)^\star$ instead of $\mathcal{P}_f(A \times X)$ this question does not occur, because such transitions are different components of an ordered sequence. Actually, an appropriate space of observations for transition systems of the form $X \to \mathcal{P}_f(A \times X)$ can be constructed by quotienting the space of observations of transition systems $X \to (A \times X)^\star$ with respect to bisimulation, see [8] for categorical details, or also [64] for some more explanations. Bisimulation with respect to this functor will be described in Example 7.1.

What is left is the general situation of arbitrary, not necessarily finitely non-deterministic, transition systems $X \to \mathcal{P}(A \times X)$. As we shall see below, the spaces of observations that we are

describing are final coalgebras. It will be shown in general that such coalgebras are isomorphisms $X \cong \mathcal{P}(A \times X)$. An easy cardinality argument shows that such an isomorphism cannot exist (for non-empty sets A). Hence there is no such space of observations ... in the world of ordinary sets. However, in the world of non-well-founded sets and classes, such a space does exist, and it plays an important role in giving meaning to various process operators, see [1, 2, 64, 67].

(vi) In another variation on the second example we may consider a machine with two buttons $(\mathsf{value}, \mathsf{next}) \colon X \to A \times X$, as above, but with autonomous activity in time. What we mean is that the machine may perform actions (and thereby change its state) if time progresses, without someone pressing a button. Hence, if at some stage the machine is in state $s \in X$, then after some time interval of length $\alpha \in \mathbb{R}_{\geq 0} = \{x \in \mathbb{R} \mid x \geq 0\}$ in which no buttons have been pressed, the machine is possibly in a different state $s' \in X$ (depending[2] on α). Such an action of time on states in combination with (coalgebraic) operations $\mathsf{value}, \mathsf{next}$ acting on states, may be used in abstract descriptions of hybrid systems, combining discrete and continuous behaviour. The possible observations that we can make in this situation are of the form: after pushing the next button $n \in \mathbb{N}$ times with intervals of length $\alpha_1, \ldots, \alpha_n$, we see a certain value in A after an interval of length β. This yields a space of observations

$$A^{(\mathbb{R}_{\geq 0})^\star \times \mathbb{R}_{\geq 0}} = A^{(\mathbb{R}_{\geq 0})^+}.$$

And the associated behaviour function $\mathsf{beh} \colon X \to A^{(\mathbb{R}_{\geq 0})^+}$ forms a suitable homomorphism as in (2).

(vii) Another example of the pattern given by the finality diagram (2) involves the definition of Böhm-trees for (untyped) lambda-terms, see [7, 10.1.3 and 10.1.4]. Briefly, the Böhm-tree $\mathrm{BT}(M)$ of a λ-term M is obtained as follows. Find out if M has a head normal form. If not, then $\mathrm{BT}(M)$ consists of a single, unlabeled node. If M does have a normal form, say of the form $\lambda \vec{x}.\, y M_1 \cdots M_n$, put

$$\mathrm{BT}(M) = \left(\begin{array}{ccccc} & & \lambda \vec{x}.y & & \\ & \swarrow & | & \searrow & \\ \mathrm{BT}(M_1) & & \cdots & & \mathrm{BT}(M_n) \end{array} \right)$$

Such a tree can be seen as arising from observations about the λ-term M. What can be observed about such a term are its abstraction variables and head variable (if any). The operation $\mathrm{BT}(-)$ of taking Böhm trees can then be seen as a function like in (2) from the set of lambda terms (modulo β-equivalence) to an appropriate space of observations (consisting of possibly infinite trees labeled by sequences of variables with finitely many branches)[3].

(viii) Other examples of interesting coalgebras can be obtained by supplying the underlying set of states with additional structure. For instance, a one-dimensional discrete time dynamical systems (X, f) consist of a complete metric space X (with distance function d_X) and a continuous function $f : X \to X$, which describes the dynamical behaviour of the system. Such systems occur, for instance, in population biology and physics. One of the main themes in the theory of dynamical systems is the systematic study of *orbits*: if $x \in X$ then its orbit is the set

$$< x >= \{x,\ f^{(1)}(x),\ f^{(2)}(x),\ f^{(3)}(x),\ \ldots\},$$

where $f^{(n+1)}(x) = f(f^{(n)}(x))$. In coalgebraic terms, the orbit of x is just the smallest subsystem (i.e., subcoalgebra) of (X, f) that contains x. Questions to be addressed are, for instance, whether a point x is periodic ($x = f^{(n)}(x)$, for some $n \geq 0$); whether there are many such periodic points and how they are distributed over X (e.g., do they form a dense subset?); and whether orbits

[2] The canonical way to describe a functional dependence on α is via a so-called monoid action $\mu \colon X \times \mathbb{R}_{\geq 0} \to X$ satisfying $\mu(s, 0) = s$ and $\mu(s, \alpha + \beta) = \mu(\mu(s, \alpha), \beta)$, see [36].

[3] In personal communication, the author of [7], Henk Barendregt, said that at the time of writing this definition of Böhm trees he felt slightly uncomfortable about the nature of the definition. He saw it as a possibly infinite process, which is well-defined at every finite stage. He emphasised, see *loc. cit.* Remark (ii) on p. 216, that it is certainly not an inductive definition. Indeed, it is a coinductive definition!

$< x >$ and $< y >$ are similar if we know that x and y are close, that is, $d_X(x, y)$ is small. As it turns out, at least one important technique that is used in the world of dynamical systems to answer some of such questions, called *symbolic dynamics* (cf. [11]), can be described elegantly in the theory of coalgebras, using amongst others techniques from metric domain theory [3, 63]. The interested reader is referred to [65] for further details, which are outside the scope of the present tutorial.

In this series of examples of coalgebras we see each time a state space X about which we make no assumptions. On this state space a function is defined, often consisting of different components, which allow us either to observe some aspect of the state space directly, or to move on to next states. We have limited access to this state space in the sense that we can only observe or modify it via these specified operations. In such a situation all that we can describe about a particular state is its behaviour, which arises by making successive observations. This will lead to the notion of bisimilarity of states: it expresses of two states that we cannot distinguish them via the operations that are at our disposal, i.e. that they are "equal as far as we can see". But this does not mean that these states are also identical as elements of X. Bisimilarity is an important, and typically coalgebraic, concept.

The above examples are meant to suggest the difference between construction in algebra, and observation in coalgebra. This difference will be described more formally below. In practice it is not always straightforward to distinguish between algebraic and coalgebraic aspects, for the following two reasons.

1. Certain abstract operations, like $X \times A \to X$, can be seen as both algebraic and coalgebraic. Algebraically, such an operation allows us to build new elements in X starting from given elements in X and parameters in A. Coalgebraically, this operation is often presented in the equivalent from $X \to X^A$ using function types. It is then seen as acting on the state space X, and yielding for each state a function from A to X which produces for each parameter element in A a next state. The context should make clear which view is prevalent. But operations of the form $A \to X$ are definitely algebraic (because they gives us information about how to put elements in X), and operations of the form $X \to A$ are coalgebraic (because they give us observable attribute values holding for elements of X). A further complication at this point is that on an initial algebra X one may have operations of the form $X \to A$, obtained by initiality. An example is the length function on lists. Such operations are derived, and are not an integral part of the (definition of the) algebra. Dually, one may have derived operations $A \to X$ on a final coalgebra X.

2. Algebraic and coalgebraic structures may be found in different hierarchic layers. For example, one can start with certain algebras describing one's application domain. On top of these one can have certain dynamical systems (processes) as coalgebras, involving such algebras (e.g. as codomains of attributes). And such coalgebraic systems may exist in an algebra of processes.

A concrete example of such layering of coalgebra on top of algebra is given by Plotkin's so-called structural operational semantics [60]. It involves a transition system (a coalgebra) describing the operational semantics of some language, by giving the transition rules by induction on the structure of the terms of the language. The latter means that the set of terms of the language is used as (initial) algebra. See [64, 67] for an investigation of this perspective. Hidden sorted algebras, see [23, 22, 14, 24, 44] can be seen as other examples: they involve "algebras" with "invisible" sorts, playing a (coalgebraic) role of a state space. Coinduction is used to reason about such hidden state spaces, see [24].

3 Inductive and coinductive definitions

In the previous section we have seen that "constructor" and "destructor/observer" operations play an important role for algebras and coalgebras, respectively. Constructors tell us how to generate

our (algebraic) data elements: the empty list constructor nil and the prefix operation cons generate all finite lists. And destructors (or observers, or transition functions) tell us what we can observe about our data elements: the head and tail operations tell us all about infinite lists: head gives a direct observation, and tail returns a next state.

Once we are aware of this duality between constructing and observing, it is easy to see the difference between inductive and coinductive definitions (relative to given collections of constructors and destructors):

> In an *inductive definition* of a function f,
> one defines the value of f on all constructors.

And:

> In a *coinductive definition* of a function f,
> one defines the values of all destructors on each outcome $f(x)$.

Such a coinductive definition determines the observable behaviour of each $f(x)$.

We shall illustrate inductive and coinductive definitions in some examples involving finite lists (with constructors nil and cons) and infinite lists (with destructors head and tail) over a fixed dataset A, as in the previous section. We assume that inductive definitions are well-known, so we only mention two trivial examples: the (earlier mentioned) function len from finite lists to natural numbers giving the length, and the function empty? from finite lists to booleans $\{\mathsf{true}, \mathsf{false}\}$ telling whether a list is empty or not:

$$\begin{cases} \mathsf{len}(\mathsf{nil}) = 0 \\ \mathsf{len}(\mathsf{cons}(a,\sigma)) = 1 + \mathsf{len}(\sigma). \end{cases} \qquad \begin{cases} \mathsf{empty?}(\mathsf{nil}) = \mathsf{true} \\ \mathsf{empty?}(\mathsf{cons}(a,\sigma)) = \mathsf{false}. \end{cases}$$

Typically in such inductive definitions, the constructors on the left hand side appear "inside" the function that we are defining.

We turn to examples of coinductive definitions (on infinite lists, say of type A). If we have a function $f\colon A \to A$, then we would like to define an extension $\mathsf{ext}(f)$ of f mapping an infinite list to an infinite list by applying f componentwise. According to the above coinductive definition scheme we have to give the values of the destructors head and tail for a sequence $\mathsf{ext}(f)(\sigma)$. They should be:

$$\begin{cases} \mathsf{head}(\mathsf{ext}(f)(\sigma)) = f(\mathsf{head}(\sigma)) \\ \mathsf{tail}(\mathsf{ext}(f)(\sigma)) = \mathsf{ext}(f)(\mathsf{tail}(\sigma)) \end{cases}$$

Here we clearly see that on the left hand side, the function that we are defining occurs "inside" the destructors. At this stage it is not yet clear if $\mathsf{ext}(f)$ is well-defined, but this is not our concern at the moment.

Alternatively, using the transition relation notation from Example (iv) in the previous section, we can write the definition of $\mathsf{ext}(f)$ as:

$$\frac{\sigma \xrightarrow{a} \sigma'}{\mathsf{ext}(f)(\sigma) \xrightarrow{f(a)} \mathsf{ext}(f)(\sigma')}$$

Suppose next, that we wish to define an operation odd which takes an infinite list, and produces a new infinite list which contains (in order) all the elements occurring in oddly numbered places of the original list. A little thought leads to the following definition clauses.

$$\begin{cases} \mathsf{head}(\mathsf{odd}(\sigma)) = \mathsf{head}(\sigma) \\ \mathsf{tail}(\mathsf{odd}(\sigma)) = \mathsf{odd}(\mathsf{tail}(\mathsf{tail}(\sigma))) \end{cases}$$

Or, in the transition relation notation:

$$\frac{\sigma \xrightarrow{a} \sigma' \xrightarrow{a'} \sigma''}{\mathsf{odd}(\sigma) \xrightarrow{a} \mathsf{odd}(\sigma'')}$$

Let us convince ourselves that this definition gives us what we want. The first clause says that the first element of the list $\mathsf{odd}(\sigma)$ is the first element of σ. The next element in $\mathsf{odd}(\sigma)$ is $\mathsf{head}(\mathsf{tail}(\mathsf{odd}(\sigma)))$, and can be computed as

$$\mathsf{head}(\mathsf{tail}(\mathsf{odd}(\sigma))) = \mathsf{head}(\mathsf{odd}(\mathsf{tail}(\mathsf{tail}(\sigma)))) = \mathsf{head}(\mathsf{tail}(\mathsf{tail}(\sigma))).$$

Hence the second element in $\mathsf{odd}(\sigma)$ is the third element in σ. It is not hard to show for $n \in \mathbb{N}$ that $\mathsf{head}(\mathsf{tail}^{(n)}(\mathsf{odd}(\sigma)))$ is the same as $\mathsf{head}(\mathsf{tail}^{(2n)}(\sigma))$.

In a similar way one can coinductively define a function even which keeps all the evenly listed elements. But it is much easier to define even as: $\mathsf{even} = \mathsf{odd} \circ \mathsf{tail}$.

As another example, we consider the merge of two infinite lists σ, τ into a single list, by taking elements from σ and τ in turn, starting with σ, say. A coinductive definition of such a function merge requires the outcomes of the destructors head and tail on $\mathsf{merge}(\sigma, \tau)$. They are given as:

$$\begin{cases} \mathsf{head}(\mathsf{merge}(\sigma, \tau)) = \mathsf{head}(\sigma) \\ \mathsf{tail}(\mathsf{merge}(\sigma, \tau)) = \mathsf{merge}(\tau, \mathsf{tail}(\sigma)) \end{cases}$$

In transition system notation, this definition looks as follows.

$$\frac{\sigma \xrightarrow{a} \sigma'}{\mathsf{merge}(\sigma, \tau) \xrightarrow{a} \mathsf{merge}(\tau, \sigma')}$$

Now one can show that the n-th element of σ occurs as $2n$-th element in $\mathsf{merge}(\sigma, \tau)$, and that the n-th element of τ occurs as $(2n+1)$-th element of $\mathsf{merge}(\sigma, \tau)$:

$$\begin{aligned} \mathsf{head}(\mathsf{tail}^{(2n)}(\sigma, \tau))) &= \mathsf{head}(\mathsf{tail}^{(n)}(\sigma)) \\ \mathsf{head}(\mathsf{tail}^{(2n+1)}(\sigma, \tau))) &= \mathsf{head}(\mathsf{tail}^{(n)}(\tau)). \end{aligned}$$

One can also define a function $\mathsf{merge}_{2,1}(\sigma, \tau)$ which takes two elements of σ for every element of τ. We leave this function as an exercise to the interested reader[4].

An obvious result that we would like to prove is: merging the lists of oddly and evenly occuring elements in a list σ returns the original list σ. That is: $\mathsf{merge}(\mathsf{odd}(\sigma), \mathsf{even}(\sigma)) = \sigma$. From what we have seen above we can easily compute that the n-th elements on both sides are equal:

$$\begin{aligned} \mathsf{head}(\mathsf{tail}^{(n)}(\mathsf{merge}(\mathsf{odd}(\sigma), \mathsf{even}(\sigma)))) &= \begin{cases} \mathsf{head}(\mathsf{tail}^{(m)}(\mathsf{odd}(\sigma))) & \text{if } n = 2m \\ \mathsf{head}(\mathsf{tail}^{(m)}(\mathsf{even}(\sigma))) & \text{if } n = 2m+1 \end{cases} \\ &= \begin{cases} \mathsf{head}(\mathsf{tail}^{(2m)}(\sigma)) & \text{if } n = 2m \\ \mathsf{head}(\mathsf{tail}^{(2m+1)}(\sigma)) & \text{if } n = 2m+1 \end{cases} \\ &= \mathsf{head}(\mathsf{tail}^{(n)}(\sigma)). \end{aligned}$$

There is however a more elegant coinductive proof-technique, which will be presented later: in Example 6.3 using uniqueness—based on the finality diagram (2)—and in the beginning of Section 7 using bisimulations.

4 Functoriality of products, coproducts and powersets

In the remainder of this paper we shall put the things we have discussed so far in a general framework. Doing so properly requires a certain amount of category theory. We do not intend to describe the relevant matters at the highest level of abstraction, making full use of category theory. Instead, we shall work mainly with ordinary sets. That is, we shall work in the universe given by the category of sets and functions. What we do need is that many operations on sets are "functorial".

[4] Stop reading here if you do not want a hint ... First, give a coinductive definition of a function $\mathsf{merge3}(\sigma, \tau, \rho)$ which merges three infinite lists in a round robin way.

This means that they do not act only on sets, but also on functions between sets, in an appropriate manner. This is familiar in the computer science literature, not in categorical terminology, but using a "map" terminology. For example, if $\mathsf{list}(A) = A^\star$ describes the set of finite lists of elements of a set A, then for a function $f\colon A \to B$ one can define a function $\mathsf{list}(A) \to \mathsf{list}(B)$ between the corresponding sets of lists, which is usually called[5] $\mathsf{map\_list}(f)$. It sends a finite list $(a_1, \ldots, a_n)$ of elements of A to the list $(f(a_1), \ldots, f(a_n))$ of elements of B, by applying f elementwise. It is not hard to show that this $\mathsf{map\_list}$ operation preserves identity functions and composite functions, i.e. that $\mathsf{map\_list}(id_A) = id_{\mathsf{list}(A)}$ and $\mathsf{map\_list}(g \circ f) = \mathsf{map\_list}(g) \circ \mathsf{map\_list}(f)$. This preservation of identities and compositions is the appropriateness that we mentioned above. In this section we concentrate on such functoriality of several basic operations, such as products, coproducts (disjoint unions) and powersets. It will be used in later sections.

We recall that for two sets X, Y the Cartesian product $X \times Y$ is the set of pairs

$$X \times Y = \{(x, y) \mid x \in X \text{ and } y \in Y\}.$$

There are then obvious projection functions $\pi\colon X \times Y \to X$ and $\pi'\colon X \times Y \to Y$ by $\pi(x, y) = x$ and $\pi'(x, y) = y$. Also, for functions $f\colon Z \to X$ and $g\colon Z \to Y$ there is a unique "pair function" $\langle f, g\rangle\colon Z \to X \times Y$ with $\pi \circ \langle f, g\rangle = f$ and $\pi' \circ \langle f, g\rangle = g$, namely $\langle f, g\rangle(z) = (f(z), g(z)) \in X \times Y$ for $z \in Z$. Notice that $\langle \pi, \pi'\rangle = id\colon X \times Y \to X \times Y$ and that $\langle f, g\rangle \circ h = \langle f \circ h, g \circ h\rangle\colon W \to X \times Y$, for functions $h\colon W \to Z$.

Interestingly, the product operation $(X, Y) \mapsto X \times Y$ does not only apply to sets, but also to functions: for functions $f\colon X \to X'$ and $g\colon Y \to Y'$ we can define a function $X \times X' \to Y \times Y'$ by $(x, y) \mapsto (f(x), g(y))$. One writes this function as $f \times g\colon X \times Y \to X' \times Y'$, whereby the symbol $\times$ is overloaded: it is used both on sets and on functions. We note that $f \times g$ can be described in terms of projections and pairing as $f \times g = \langle f \circ \pi, g \circ \pi'\rangle$. It is easily verified that the operation $\times$ on functions satisfies

$$id_X \times id_Y = id_{X \times Y} \qquad \text{and} \qquad (f \circ h) \times (g \circ k) = (f \times g) \circ (h \times k).$$

This expresses that the product $\times$ is *functorial*: it does not only apply to sets, but also to functions; and it does so in such a way that identity maps and composites are preserved.

Many more operations are functorial. Also the coproduct (or disjoint union, or sum) $+$ is. For sets X, Y we write their disjoint union as $X + Y$. Explicitly:

$$X + Y = \{\langle 0, x\rangle \mid x \in X\} \cup \{\langle 1, y\rangle \mid y \in Y\}.$$

The first components 0 and 1 serve to force this union to be disjoint. These "tags" enables us to recognise the elements of X and of Y inside $X + Y$. Instead of projections as above we now have "coprojections" $\kappa\colon X \to X + Y$ and $\kappa'\colon Y \to X + Y$ going in the other direction. One puts $\kappa(x) = \langle 0, x\rangle$ and $\kappa'(y) = \langle 1, y\rangle$. And instead of tupleing we now have "cotupleing" (sometimes called "source tupleing"): for functions $f\colon X \to Z$ and $g\colon Y \to Z$ there is a unique function $[f, g]\colon X + Y \to Z$ with $[f, g] \circ \kappa = f$ and $[f, g] \circ \kappa' = g$. One defines $[f, g]$ by case distinction:

$$[f, g](w) = \begin{cases} f(x) & \text{if } w = \langle 0, x\rangle \\ g(y) & \text{if } w = \langle 1, y\rangle. \end{cases}$$

Notice that $[\kappa, \kappa'] = id$ and $h \circ [f, g] = [h \circ f, h \circ g]$.

This is the coproduct $X + Y$ on sets. We can extend it to functions in the following way. For $f\colon X \to X'$ and $g\colon Y \to Y'$ there is a function $f + g\colon X + Y \to X' + Y'$ by

$$(f + g)(w) = \begin{cases} \langle 0, f(x)\rangle & \text{if } w = \langle 0, x\rangle \\ \langle 1, g(y)\rangle & \text{if } w = \langle 1, y\rangle. \end{cases}$$

[5] In the category theory literature one uses the same name for the actions of a functor on objects and on morphisms; this leads to the notation $\mathsf{list}(f)$ or $f^\star$ for this function $\mathsf{map\_list}(f)$.

Equivalently, we could have defined: $f+g = [\kappa \circ f, \kappa' \circ g]$. This operation $+$ on functions preserves identities and composition:

$$id_X + id_Y = id_{X+Y} \qquad \text{and} \qquad (f \circ h) + (g \circ k) = (f+g) \circ (h+k).$$

We should emphasise that this coproduct $+$ is very different from ordinary union $\cup$. For example, $\cup$ is idempotent: $X \cup X = X$, but there is not even an isomorphism between $X + X$ and X (if $X \neq \emptyset$).

For a fixed set A, the assignment $X \mapsto X^A = \{f \mid f \text{ is a function } A \to X\}$ is functorial: a function $g\colon X \to Y$ yields a function $g^A\colon X^A \to Y^A$ sending $f \in X^A$ to $(g \circ f) \in Y^A$. Clearly, $id^A = id$ and $(h \circ g)^A = h^A \circ g^A$.

Another example of a functorial operation is powerset: $X \mapsto \mathcal{P}(X)$. For a function $f\colon X \to X'$ one defines $\mathcal{P}(f)\colon \mathcal{P}(X) \to \mathcal{P}(X')$ by

$$U \mapsto \{f(x) \mid x \in U\}.$$

Then $\mathcal{P}(id_X) = id_{\mathcal{P}(X)}$ and $\mathcal{P}(f \circ h) = \mathcal{P}(f) \circ \mathcal{P}(h)$. We shall write $\mathcal{P}_{\mathrm{fin}}(-)$ for the (functorial) operation which maps X to the set of its *finite* subsets.

Here are some trivial examples of functors. The identity operation $X \mapsto X$ is functorial: it acts on functions as $f \mapsto f$. And for a constant set C we have a constant functorial operation $X \mapsto C$; a function $f\colon X \to X'$ is mapped to the identity function $id_C\colon C \to C$.

Once we know these actions on functions, we can define functorial operations (or: *functors*, for short) merely by giving their actions on sets. We will often say things like: consider the functor

$$T(X) = X + (C \times X).$$

The action on sets is then $X \mapsto X + (C \times X)$. And for a function $f\colon X \to X'$ we have an action $T(f)$ of the functor T on f as a function $T(f)\colon T(X) \to T(X')$. Explicitly, $T(f)$ is the function

$$f + (id_C \times f)\colon X + (C \times X) \longrightarrow X' + (C \times X')$$

given by:

$$w \mapsto \begin{cases} \langle 0, f(x)\rangle & \text{if } w = \langle 0, x\rangle \\ \langle 1, (c, f(x))\rangle & \text{if } w = \langle 1, (c, x)\rangle. \end{cases}$$

In the sequel we shall only use such "polynomial" functors T that are built up with constants, identity functors, products, coproducts and also (finite) powersets. We describe these functors by only giving their actions on sets.

(There is a more general notion of 'functor' as mapping from one "category" to another. Here we are only interested in these polynomial functors, going from the category of sets and functions to itself. But much of the theory applies to more general situations.)

We shall write $1 = \{*\}$ for a singleton set, with typical inhabitant $*$. Notice that for every set X there is precisely one function $X \to 1$. This says that 1 is *final* (or *terminal*) in the category of sets and functions. And functions $1 \to X$ correspond to elements of X. Usually we shall identify the two. We write 0 for the empty set. For every set X there is precisely one function $0 \to X$, namely the empty function. This property is the *initiality* of 0. (These sets 1 and 0 can be seen as the empty product and coproduct.)

We list some useful isomorphisms.

$$\begin{array}{rcl@{\qquad}rcl} X \times Y & \cong & Y \times X & X + Y & \cong & Y + X \\ 1 \times X & \cong & X & 0 + X & \cong & X \\ X \times (Y \times Z) & \cong & (X \times Y) \times Z & X + (Y + Z) & \cong & (X + Y) + Z \\ X \times 0 & \cong & 0 & X \times (Y + Z) & \cong & (X \times Y) + (X \times Z). \end{array}$$

The last two isomorphisms describe the distribution of products over finite coproducts. We shall often work "up-to" the above isomorphisms, so that we can simply write an n-ary product as $X_1 \times \cdots \times X_n$ without bothering about bracketing.

5 Algebras and induction

In this section we start by showing how polynomial functors—as introduced in the previous section—can be used to describe signatures of operations. Algebras of such functors correspond to models of such signatures. They consist of a carrier set with certain functions interpreting the operations. A general notion of homomorphism is defined between such algebras of a functor. This allows us to define initial algebras by the following property: for an arbitrary algebra there is precisely one homomorphism from the initial algebra to this algebra. This turns out to be a powerful notion. It captures algebraic structures which are generated by constructor operations, as will be shown in several examples. Also, it gives rise to the familiar principles of definition by induction and proof by induction.

We start with an example. Let T be the polynomial functor $T(X) = 1 + X + (X \times X)$, and consider for a set U a function $a\colon T(U) \to U$. Such a map a may be identified with a 3-cotuple $[a_1, a_2, a_3]$ of maps $a_1\colon 1 \to U$, $a_2\colon U \to U$ and $a_3\colon U \times U \to U$ giving us three separate functions going into the set U. They form an example of an algebra (of the functor T): a set together with a (cotupled) number of functions going into that set. For example, if one has a group G, with unit element $e\colon 1 \to G$, inverse function $i\colon G \to G$ and multiplication function $m\colon G \times G \to G$, then one can organise these three maps as an algebra $[e, i, m]\colon T(G) \to G$ via cotupling[6]. The shape of the functor T determines a certain signature of operations. Had we taken a different functor $S(X) = 1 + (X \times X)$, then maps (algebras of S) $S(U) \to U$ would capture pairs of functions $1 \to U$, $U \times U \to U$ (e.g. of a monoid).

Definition 5.1 Let T be a functor. An *algebra* of T (or, a T-*algebra*) is a pair consisting of a set U and a function $a\colon T(U) \to U$.

We shall call the set U the *carrier* of the algebra, and the function a the *algebra structure*, or also the *operation* of the algebra.

For example, the zero and successor functions $0\colon 1 \to \mathbb{N}$, $S\colon \mathbb{N} \to \mathbb{N}$ on the natural numbers form an algebra $[0, S]\colon 1 + \mathbb{N} \to \mathbb{N}$ of the functor $T(X) = 1 + X$. And the set of A-labeled finite binary trees $\mathsf{Tree}(A)$ comes with functions $\mathsf{nil}\colon 1 \to \mathsf{Tree}(A)$ for the empty tree, and $\mathsf{node}\colon \mathsf{Tree}(A) \times A \times \mathsf{Tree}(A) \to \mathsf{Tree}(A)$ for constructing a tree out of two (sub)trees and a (node) label. Together, nil and node form an algebra $1 + (\mathsf{Tree}(A) \times A \times \mathsf{Tree}(A)) \to \mathsf{Tree}(A)$ of the functor $S(X) = 1 + (X \times A \times X)$.

We illustrate the link between signatures (of operations) and functors with further details. Let Σ be a (single-sorted, or single-typed) signature, given by a finite collection Σ of operations σ, each with an arity $\mathrm{ar}(\sigma) \in \mathbb{N}$. Each $\sigma \in \Sigma$ will be understood as an operation

$$\sigma\colon \underbrace{X \times \cdots \times X}_{\mathrm{ar}(\sigma)\text{ times}} \longrightarrow X$$

taking $\mathrm{ar}(\sigma)$ inputs of some type X, and producing an output of type X. With this signature Σ, say with set of operations $\{\sigma_1, \ldots, \sigma_n\}$ we associate a functor

$$T_\Sigma(X) = X^{\mathrm{ar}(\sigma_1)} + \cdots + X^{\mathrm{ar}(\sigma_n)},$$

where for $m \in \mathbb{N}$ the set X^m is the m-fold product $X \times \cdots \times X$. An algebra $a\colon T_\Sigma(U) \to U$ of this functor T_Σ can be identified with an n-cotuple $a = [a_1, \ldots a_n]\colon U^{\mathrm{ar}(\sigma_1)} + \cdots + U^{\mathrm{ar}(\sigma_n)} \to U$ of functions $a_i\colon U^{\mathrm{ar}(\sigma_i)} \to U$ interpreting the operations σ_i in Σ as functions on U. Hence algebras of the functor T_Σ correspond to models of the signature Σ. One sees how the arities in the signature Σ determine the shape of the associated functor T_Σ. Notice that as special case when an arity of an operation is zero we have a constant in Σ. In a T_Σ-algebra $T_\Sigma(U) \to U$ we get an associated map $U^0 = 1 \to U$ giving us an element of the carrier set U as interpretation of the constant. The assumption that the signature Σ is finite is not essential for the correspondence between models

[6] Only the group's operations, and not its equations, are captured in this map $T(G) \to G$.

of Σ and algebras of T_Σ; if Σ is infinite, one can define T_Σ via an infinite coproduct, commonly written as $T_\Sigma(X) = \coprod_{\sigma\in\Sigma} X^{\mathbf{ar}(\sigma)}$.

Polynomial functors T built up from the identity functor, products and coproducts (without constants) have algebras which are models of the kind of signatures Σ described above. This is because by the distribution of products over coproducts one can always write such a functor in "disjunctive normal form" as $T(X) = X^{m_1} + \cdots + X^{m_n}$ for certain natural numbers n and $m_1, \ldots, m_n$. The essential role of the coproducts is to combine multiple operations into a single operation.

The polynomial functors that we use are not only of this form $T(X) = X^{m_1} + \cdots + X^{m_n}$, but may also involve constant sets. This is quite useful, for example, to describe for an arbitrary set A a signature for lists of A's, with function symbols $\mathsf{nil}\colon 1 \to X$ for the empty list, and $\mathsf{cons}\colon A \times X \to X$ for prefixing an element of type A to a list. A model (interpretation) for such a signature is an algebra $T(U) \to U$ of the functor $T(X) = 1 + (A \times X)$ associated with this signature.

We turn to "homomorphisms of algebras", to be understood as structure preserving functions between algebras (of the same signature, or functor). Such a homomorphism is a function between the carrier sets of the algebras which commutes with the operations. For example, suppose we have two algebras $\ell_1\colon 1 \to U_1$, $c_1\colon A \times U_1 \to U_1$ and $\ell_2\colon 1 \to U_2$, $c_2\colon A \times U_2 \to U_2$ of the above list signature. A homomorphism of algebras from the first to the second consists of a function $f\colon U_1 \to U_2$ between the carriers with $f \circ \ell_1 = \ell_2$ and $f \circ c_1 = c_2 \circ (id \times f)$. In two diagrams:

$$\begin{array}{ccc} 1 & = & 1 \\ {\scriptstyle \ell_1}\downarrow & & \downarrow{\scriptstyle \ell_2} \\ U_1 & \xrightarrow[f]{} & U_2 \end{array} \qquad \text{and} \qquad \begin{array}{ccc} A \times U_1 & \xrightarrow{id \times f} & A \times U_2 \\ {\scriptstyle c_1}\downarrow & & \downarrow{\scriptstyle c_2} \\ U_1 & \xrightarrow[f]{} & U_2 \end{array}$$

Thus, writing $n_1 = \ell_1(*)$ and $n_2 = \ell_2(*)$, these diagrams express that $f(n_1) = n_2$ and $f(c_1(a, x)) = c_2(a, f(x))$, for $a \in A$ and $x \in U_1$.

These two diagrams can be combined into a single diagram:

$$\begin{array}{ccc} 1 + (A \times U_1) & \xrightarrow{id + (id \times f)} & 1 + (A \times U_2) \\ {\scriptstyle [\ell_1, c_1]}\downarrow & & \downarrow{\scriptstyle [\ell_2, c_2]} \\ U_1 & \xrightarrow[f]{} & U_2 \end{array}$$

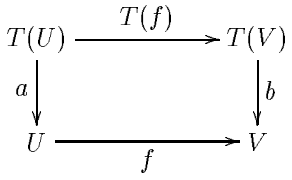
i.e., for the list-functor $T(X) = 1 + (A \times X)$,

$$\begin{array}{ccc} T(U_1) & \xrightarrow{T(f)} & T(U_2) \\ {\scriptstyle [\ell_1, c_1]}\downarrow & & \downarrow{\scriptstyle [\ell_2, c_2]} \\ U_1 & \xrightarrow[f]{} & U_2 \end{array}$$

The latter formulation is entirely in terms of the functor involved. This motivates the following definition.

Definition 5.2 Let T be a functor with algebras $a\colon T(U) \to U$ and $b\colon T(V) \to V$. A *homomorphism of algebras* (also called a *map of algebras*, or an *algebra map*) from (U, a) to (V, b) is a function $f\colon U \to V$ between the carrier sets which commutes with the operations: $f \circ a = b \circ T(f)$

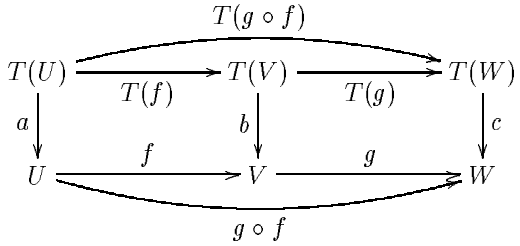
in



As a triviality we notice that for an algebra $a: T(U) \to U$ the identity function $U \to U$ is an algebra map $(U, a) \to (U, a)$. And we can compose algebra maps as functions: given two algebra maps

$$\Big(T(U) \xrightarrow{a} U\Big) \xrightarrow{f} \Big(T(V) \xrightarrow{b} V\Big) \xrightarrow{g} \Big(T(W) \xrightarrow{c} W\Big)$$

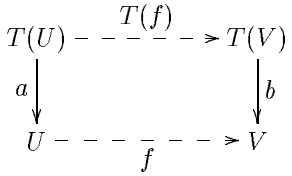
then the composite function $g \circ f: U \to W$ is an algebra map from (U, a) to (W, c). This is because $g \circ f \circ a = g \circ b \circ T(f) = c \circ T(g) \circ T(f) = c \circ T(g \circ f)$, see the following diagram.



(Thus: algebras and their homomorphisms form a category.)

Now that we have a notion of homomorphism of algebras we can formulate the important concept of "initiality" for algebras.

Definition 5.3 An algebra $a: T(U) \to U$ of a functor T is *initial* if for each algebra $b: T(V) \to V$ there is a unique homomorphism of algebras from (U, a) to (V, b). Diagrammatically we express this uniqueness by a dashed arrow, call it f, in



We shall sometimes call this f the "unique mediating algebra map".

We emphasise that unique existence has two aspects, namely *existence* of an algebra map out of the initial algebra to another algebra, and *uniqueness*, in the form of equality of any two algebra maps going out of the initial algebra to some other algebra. Existence will be used as an (inductive) definition principle, and uniqueness as an (inductive) proof principle.

As a first example, we shall describe the set $\mathbb{N}$ of natural numbers as initial algebra.

Example 5.4 Consider the set $\mathbb{N}$ of natural number with its zero and successor function $0: 1 \to \mathbb{N}$ and $S: \mathbb{N} \to \mathbb{N}$. These functions combine into a single function $[0, S]: 1 + \mathbb{N} \to \mathbb{N}$, forming an algebra of the functor $T(X) = 1 + X$. We will show that this map $[0, S]: 1 + \mathbb{N} \to \mathbb{N}$ is the initial algebra of this functor. And this characterises the set of natural numbers (up-to-isomorphism), by Lemma 5.5 (ii) below.

To prove initiality, assume we have an arbitrary set U carrying a T-algebra structure $[u, h]: 1 + U \to U$. We have to define a "mediating" homomorphism $f: \mathbb{N} \to U$. We try iteration:

$$f(n) = h^{(n)}(u)$$

(where we simply write u instead of $u(*)$). That is,

$$f(0) = u \qquad \text{and} \qquad f(n+1) = h(f(n)).$$

These two equations express that we have a commuting diagram

$$\begin{array}{ccc} 1+\mathbb{N} & \xrightarrow{id+f} & 1+U \\ {\scriptstyle [0,S]}\downarrow & & \downarrow{\scriptstyle [u,h]} \\ \mathbb{N} & \xrightarrow[f]{} & U \end{array}$$

making f a homomorphism of algebras. This can be verified easily by distinguishing for an arbitrary element $x \in 1+\mathbb{N}$ in the upper-left corner the two cases $x = (0,*) = \kappa(*)$ and $x = (1,n) = \kappa'(n)$, for $n \in \mathbb{N}$. In the first case $x = \kappa(*)$ we get

$$f([0,S](\kappa(*))) = f(0) = u = [u,h](\kappa(*)) = [u,h]((id+f)(\kappa(*))).$$

In the second case $x = \kappa'(n)$ we similarly check:

$$f([0,S](\kappa'(n))) = f(S(n)) = h(f(n)) = [u,h](\kappa'(f(n))) = [u,h]((id+f)(\kappa'(n))).$$

Hence we may conclude that $f([0,S](x)) = [u,h]((id+f)(x))$, for all $x \in 1+\mathbb{N}$, i.e. that $f \circ [0,S] = [u,h] \circ (id+f)$.

This looks promising, but we still have to show that f is the only map making the diagram commute. If $g\colon \mathbb{N} \to U$ also satisfies $g \circ [0,S] = [u,h] \circ (id+g)$, then $g(0) = u$ and $g(n+1) = h(g(n))$, by the same line of reasoning followed above. Hence $g(n) = f(n)$ by induction on n, so that $g = f\colon \mathbb{N} \to U$.

We shall give a simple example showing how to use this initiality for inductive definitions. Suppose we wish to define by induction the function $f(n) = 2^{-n}$ from the natural numbers $\mathbb{N}$ to the rational numbers $\mathbb{Q}$. Its defining equations are:

$$f(0) = 1 \qquad \text{and} \qquad f(n+1) = \tfrac{1}{2} f(n).$$

In order to define this function $f\colon \mathbb{N} \to \mathbb{Q}$ by initiality, we have to put an algebra structure $1 + \mathbb{Q} \to \mathbb{Q}$ on the set of rational numbers $\mathbb{Q}$, see the above definition. This algebra on $\mathbb{Q}$ corresponds to the right hand side of the two defining equations of f, given as two functions

$$\begin{array}{ccc} 1 & \xrightarrow{1} & \mathbb{Q} \\ * & \longmapsto & 1 \end{array} \qquad\qquad \begin{array}{ccc} \mathbb{Q} & \xrightarrow{\frac{1}{2}(-)} & \mathbb{Q} \\ x & \longmapsto & \frac{1}{2}x \end{array}$$

(where we use '1' both for the singleton set $1 = \{*\}$ and for the number $1 \in \mathbb{Q}$) which combine into a single function

$$1 + \mathbb{Q} \xrightarrow{[1, \frac{1}{2}(-)]} \mathbb{Q}$$

forming an algebra on $\mathbb{Q}$. The function $f(n) = 2^{-n}$ is then determined by initiality as the unique function making the following diagram commute.

$$\begin{array}{ccc} 1+\mathbb{N} & \xrightarrow{id+f} & 1+\mathbb{Q} \\ {\scriptstyle [0,S]}\downarrow & & \downarrow{\scriptstyle [1,\frac{1}{2}(-)]} \\ \mathbb{N} & \xrightarrow[f]{} & \mathbb{Q} \end{array}$$

This shows how initiality can be used to define functions by induction. It requires that one puts an appropriate algebra structure on the codomain (i.e. the range) of the intended function, corresponding to the induction clauses that determine the function.

We emphasise that the functor T is a parameter in Definitions 5.2 and 5.3 of "homomorphism" and "initiality" for algebras, yielding uniform notions for all functors T (representing certain signatures). It turns out that initial algebras have certain properties, which can be shown for all functors T at once. Diagrams are convenient in expressing and proving these properties, because they display information in a succinct way. And they are useful both in existence and uniqueness arguments.

Lemma 5.5 *Let T be a functor.*

(i) Initial T-algebras, if they exist, are unique, up-to-isomorphism of algebras. That is, if we have two initial algebras $a: T(U) \to U$ and $a': T(U') \to U'$ of T, then there is a unique isomorphism $f: U \xrightarrow{\cong} U'$ of algebras:

$$\begin{array}{ccc} T(U) & \xrightarrow{T(f)} & T(U') \\ {\scriptstyle a}\downarrow & \cong & \downarrow{\scriptstyle a'} \\ U & \xrightarrow[\cong]{f} & U' \end{array}$$

(ii) The operation of an initial algebras is an isomorphism: if $a: T(U) \to U$ is initial algebra, then a has an inverse $a^{-1}: U \to T(U)$.

The first point tells us that a functor can have (essentially) at most one initial algebra[7]. Therefore, we often speak of *the* initial algebra of a functor T. And the second point—which is due to Lambek—says that an initial algebra $T(U) \to U$ is a fixed point $T(U) \cong U$ of the functor T. Initial algebras may be seen as generalizations of least fixed points of monotone functions, since they have a (unique) map into an arbitrary algebra.

Proof: (i) Suppose both $a: T(U) \to U$ and $a': T(U') \to U'$ are initial algebras of the functor T. By initiality of a there is a unique algebra map $f: U \to U'$. Similarly, by initiality of a' there is a unique algebra map $f': U' \to U$ in the other direction:

$$\begin{array}{ccc} T(U) & \overset{T(f)}{\dashrightarrow} & T(U') \\ {\scriptstyle a}\downarrow & & \downarrow{\scriptstyle a'} \\ U & \underset{f}{\dashrightarrow} & U' \end{array} \qquad\qquad \begin{array}{ccc} T(U') & \overset{T(f')}{\dashrightarrow} & T(U) \\ {\scriptstyle a'}\downarrow & & \downarrow{\scriptstyle a} \\ U' & \underset{f'}{\dashrightarrow} & U \end{array}$$

Here we use the existence parts of initiality. The uniqueness part gives us that the two resulting algebra maps $(U, a) \to (U, a)$, namely $f \circ f'$ and id in:

$$\begin{array}{ccccc} T(U) & \xrightarrow{T(f)} & T(U') & \xrightarrow{T(f')} & T(U) \\ {\scriptstyle a}\downarrow & & {\scriptstyle a'}\downarrow & & \downarrow{\scriptstyle a} \\ U & \xrightarrow[f]{} & U' & \xrightarrow[f']{} & U \end{array} \qquad \text{and} \qquad \begin{array}{ccc} T(U) & \xrightarrow{T(id)} & T(U) \\ {\scriptstyle a}\downarrow & & \downarrow{\scriptstyle a} \\ U & \xrightarrow[id]{} & U \end{array}$$

must be equal, i.e. that $f' \circ f = id$. Uniqueness of algebra maps $(U', a') \to (U', a')$ similarly yields $f \circ f' = id$. Hence f is an isomorphism of algebras.

(ii) Let $a: T(U) \to U$ be initial T-algebra. In order to show that the function a is an isomorphism, we have to produce an inverse function $U \to T(U)$. Initiality of (U, a) can be used to define functions out of U to arbitrary algebras. Since we seek a function $U \to T(U)$, we have to put an algebra structure on the set $T(U)$. A moment's thought yields a candidate, namely the result

[7] This is a more general property of initial objects in a category.

$T(a)\colon T(T(U)) \to T(U)$ of applying the functor T to the function a. This function $T(a)$ gives by initiality of $a\colon T(U) \to U$ rise to a function $a'\colon U \to T(U)$ with $T(a) \circ T(a') = a' \circ a$ in:

$$\begin{array}{ccc} T(U) & \overset{T(a')}{\dashrightarrow} & T(T(U)) \\ {\scriptstyle a}\downarrow & & \downarrow{\scriptstyle T(a)} \\ U & \underset{a'}{\dashrightarrow} & T(U) \end{array}$$

The function $a \circ a'\colon U \to U$ is an algebra map $(U, a) \to (U, a)$:

$$\begin{array}{ccccc} T(U) & \xrightarrow{T(a')} & T(T(U)) & \xrightarrow{T(a)} & T(U) \\ {\scriptstyle a}\downarrow & & \downarrow{\scriptstyle T(a)} & & \downarrow{\scriptstyle a} \\ U & \xrightarrow[a']{} & T(U) & \xrightarrow[a]{} & U \end{array}$$

so that $a \circ a' = id$ by uniqueness of algebra maps $(U, a) \to (U, a)$. But then

$$\begin{array}{rcll} a' \circ a & = & T(a) \circ T(a') & \text{by definition of } a' \\ & = & T(a \circ a') & \text{since } T \text{ preserves composition} \\ & = & T(id) & \text{as we have just seen} \\ & = & id & \text{since } T \text{ preserves identities.} \end{array}$$

Hence $a\colon T(U) \to U$ is an isomorphism with a' as its inverse. □

From now on we shall often write an initial T-algebra as a map $a\colon T(U) \xrightarrow{\cong} U$, making this isomorphism explicit.

Example 5.6 Let A be fixed set and consider the functor $T(X) = 1 + (A \times X)$ that we used earlier to capture models of the list signature $1 \to X$, $A \times X \to X$. We claim that the initial algebra of T is the set $A^\star = \mathsf{list}(A) = \bigcup_{n \in \mathbb{N}} A^n$ of finite sequences of elements of A, together with the function (or element) $1 \to A^\star$ given by the empty list $\mathsf{nil} = ()$, and the function $A \times A^\star \to A^\star$ which maps an element $a \in A$ and a list $\alpha = (a_1, \ldots, a_n) \in A^\star$ to the list $\mathsf{cons}(a, \alpha) = (a, a_1, \ldots, a_n) \in A^\star$, obtained by prefixing a to α. These two functions combine into a single function $[\mathsf{nil}, \mathsf{cons}]\colon 1 + (A \times A^\star) \to A^\star$, which, as one easily checks, is an isomorphism. But this does not yet mean that it is the initial algebra. We will check this explicitly.

For an arbitrary algebra $[u, h]\colon 1 + (A \times U) \to U$ of the list-functor T we have a unique homomorphism $f\colon A^\star \to U$ of algebras:

$$\begin{array}{ccc} 1 + (A \times A^\star) & \xrightarrow{id + (id \times f)} & 1 + (A \times U) \\ {\scriptstyle [\mathsf{nil}, \mathsf{cons}]}\downarrow & & \downarrow{\scriptstyle [u, h]} \\ A^\star & \underset{f}{\longrightarrow} & U \end{array}$$

namely

$$f(\alpha) = \begin{cases} u & \text{if } \alpha = \mathsf{nil} \\ h(a, f(\beta)) & \text{if } \alpha = \mathsf{cons}(a, \beta). \end{cases}$$

We leave it to the reader to verify that f is indeed the unique function $A^\star \to U$ making the diagram commute.

Again we can use this initiality of $A^\star$ to define functions by induction (for lists). As example we take the length function $\mathsf{len}\colon A^\star \to \mathbb{N}$, described already in the beginning of Section 2. In order to define it by initiality, it has to arise from a list-algebra structure $1 + A \times \mathbb{N} \to \mathbb{N}$ on the natural numbers $\mathbb{N}$. This algebra structure is the cotuple of the two functions $0\colon 1 \to \mathbb{N}$ and $S \circ \pi'\colon A \times \mathbb{N} \to \mathbb{N}$. Hence len is determined as the unique function in the following initiality diagram.

$$\begin{array}{ccc} 1 + (A \times A^\star) & \xrightarrow{id + (id \times \mathsf{len})} & 1 + (A \times \mathbb{N}) \\ {\scriptstyle [\mathsf{nil},\mathsf{cons}]}\downarrow \cong & & \downarrow{\scriptstyle [0, S \circ \pi']} \\ A^\star & \xrightarrow[\mathsf{len}]{} & \mathbb{N} \end{array}$$

The algebra structure that we use on $\mathbb{N}$ corresponds to the defining clauses $\mathsf{len}(\mathsf{nil}) = 0$ and $\mathsf{len}(\mathsf{cons}(a,\alpha)) = S(\mathsf{len}(\alpha)) = S(\mathsf{len}(\pi'(a,\alpha))) = S(\pi'(id \times \mathsf{len})(a,\alpha))$.

We proceed with an example showing how proof by induction involves using the uniqueness of a map out of an initial algebra. Consider therefore the "doubling" function $d\colon A^\star \to A^\star$ which replaces each element a in a list α by two consecutive occurrences a, a in $d(\alpha)$. This function is defined as the unique one making the following diagram commute.

$$\begin{array}{ccc} 1 + (A \times A^\star) & \xrightarrow{id + (id \times d)} & 1 + (A \times A^\star) \\ {\scriptstyle [\mathsf{nil},\mathsf{cons}]}\downarrow \cong & & \downarrow{\scriptstyle [\mathsf{nil}, \lambda(a,\alpha).\,\mathsf{cons}(a,\mathsf{cons}(a,\alpha))]} \\ A^\star & \xrightarrow[d]{} & A^\star \end{array}$$

That is, d is defined by the induction clauses $d(\mathsf{nil}) = \mathsf{nil}$ and $d(\mathsf{cons}(a,\alpha)) = \mathsf{cons}(a,\mathsf{cons}(a,d(\alpha))$. We wish to show that the length of the list $d(\alpha)$ is twice the length of α, i.e. that

$$\mathsf{len}(d(\alpha)) = 2 \cdot \mathsf{len}(\alpha).$$

The ordinary induction proof consists of two steps:

$$\mathsf{len}(d(\mathsf{nil})) = \mathsf{len}(\mathsf{nil}) = 0 = 2 \cdot 0 = 2 \cdot \mathsf{len}(\mathsf{nil})$$

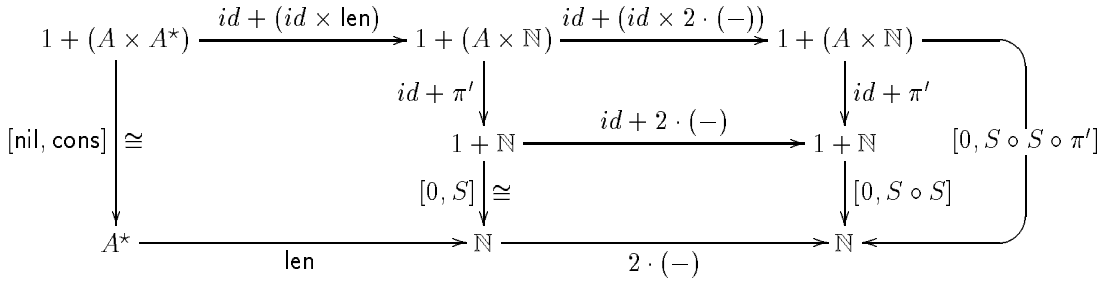
And

$$\begin{array}{rcl} \mathsf{len}(d(\mathsf{cons}(a,\alpha))) & = & \mathsf{len}(\mathsf{cons}(a,\mathsf{cons}(a,d(\alpha)))) \\ & = & 1 + 1 + \mathsf{len}(d(\alpha)) \\ & \overset{(\mathrm{IH})}{=} & 2 + 2 \cdot \mathsf{len}(\alpha) \\ & = & 2 \cdot (1 + \mathsf{len}(\alpha)) \\ & = & 2 \cdot \mathsf{len}(\mathsf{cons}(a,\alpha)). \end{array}$$

The "initiality" induction proof of the fact $\mathsf{len} \circ d = 2 \cdot (-) \circ \mathsf{len}$ uses uniqueness in the following manner. Both $\mathsf{len} \circ d$ and $2 \cdot (-) \circ \mathsf{len}$ are homomorphism from the (initial) algebra $(A^\star, [\mathsf{nil},\mathsf{cons}])$ to the algebra $(\mathbb{N}, [0, S \circ S \circ \pi'])$, so they must be equal by initiality. First we check that $\mathsf{len} \circ d$ is an appropriate homomorphism by inspection of the following diagram.

$$\begin{array}{ccccc} 1 + (A \times A^\star) & \xrightarrow{id + (id \times d)} & 1 + (A \times A^\star) & \xrightarrow{id + (id \times \mathsf{len})} & 1 + (A \times \mathbb{N}) \\ {\scriptstyle [\mathsf{nil},\mathsf{cons}]}\downarrow \cong & & \downarrow{\scriptstyle [\mathsf{nil}, \lambda(a,\alpha).\,\mathsf{cons}(a,\mathsf{cons}(a,\alpha))]} & & \downarrow{\scriptstyle [0, S \circ S \circ \pi']} \\ A^\star & \xrightarrow[d]{} & A^\star & \xrightarrow[\mathsf{len}]{} & \mathbb{N} \end{array}$$

The rectangle on the left commutes by definition of d. And commutation of the rectangle on the right follows easily from the definition of len. Next we check that $2\cdot(-)\circ\mathsf{len}$ is also a homomorphism of algebras:



The square on the left commutes by definition of len. Commutation of the upper square on the right follows from an easy computation. And the lower square on the right may be seen as defining the function $2\cdot(-)\colon\mathbb{N}\to\mathbb{N}$ by the clauses: $2\cdot 0=0$ and $2\cdot(S(n))=S(S(2\cdot n))$—which we took for granted in the earlier "ordinary" proof.

We conclude our brief discussion of algebras and induction with a few remarks.

1. Given a number of constructors one can form the carrier set of the associated initial algebra as the set of 'closed' terms (or 'ground' terms, not containing variables) that can be formed with these constructors. For example, the zero and successor constructors $\mathsf{0}\colon 1\to X$ and $\mathsf{S}\colon X\to X$ give rise to the set of closed terms,

$$\{\mathsf{0},\mathsf{S}(\mathsf{0}),\mathsf{S}(\mathsf{S}(\mathsf{0}))),\ldots\}$$

   which is (isomorphic to) the set $\mathbb{N}$ of natural numbers. Similarly, the set of closed terms arising from the A-list constructors $\mathsf{nil}\colon 1\to X$, $\mathsf{cons}\colon A\times X\to X$ is the set $A^\star$ of finite sequences (of elements of A).

   Although it is pleasant to know what an initial algebra looks like, in using initiality we do not need this knowledge. All we need to know is that there exists an initial algebra. Its defining property is sufficient to use it. There are abstract results, guaranteeing the existence of initial algebras for certain (continuous) functors, see e.g. [43, 66], where initial algebras are constructed as suitable colimits, generalizing the construction of least fixed points of continuous functions.

2. The initiality format of induction has the important advantage that it generalises smoothly from natural numbers to other (algebraic) data types, like lists or trees. Once we know the signature containing the constructor operations of these data types, we know what the associated functor is and we can determine its initial algebra. This uniformity provided by initiality was first stressed by the "ADT-group" [26], and forms the basis for inductively defined types in many programming languages. For example, in the (functional) language ML, the user can introduce a new inductive type X via the notation

$$\mathsf{datatype}\ X=c_1\ \mathsf{of}\ \sigma_1(X)\mid\cdots\mid c_n\ \mathsf{of}\ \sigma_n(X).$$

   The idea is that X is the carrier of the initial algebra associated with the constructors $c_1\colon\sigma_1(X)\to X$, …, $c_n\colon\sigma_n(X)\to X$. That is, with the functor $T(X)=\sigma_1(X)+\cdots+\sigma_n(X)$. The σ_i are existing types which may contain X (positively)[8]. The uniformity provided by the initial algebra format (and dually also by the final coalgebra format) is very useful if one

[8] This definition scheme in ML contains various aspects which are not investigated here, e.g. it allows (a) $X=X(\vec{\alpha})$ to contain type variables $\vec{\alpha}$, (b) mutual dependencies between such definitions, (c) iteration of inductive definitions (so that, for example, the LIST operation which is obtained via this scheme can be used in the σ_i.

wishes to automatically generate various rules associated with (co)inductively defined types (for example in programming languages like CHARITY [15, 16] or in proof tools like PVS [52], HOL/ISABELLE [27, 49, 55, 56], or COQ [54]).

Another great advantage of the initial algebra format is that it is dual to the final coalgebra format, as we shall see in the next section. This forms the basis for the duality between induction and coinduction.

3. We have indicated only in one example that uniqueness of maps out of an initial algebra corresponds to proof (as opposed to definition) by induction. To substantiate this claim further we show how the usual predicate formulation of induction for lists can be derived from the initial algebra formulation. This predicate formulation says that a predicate (or subset) $P \subseteq A^\star$ is equal to $A^\star$ in case $\mathsf{nil} \in P$ and $\alpha \in P \Rightarrow \mathsf{cons}(a, \alpha) \in P$ (for all $a \in A$). Let us consider P as a set in its own right, with an explicit inclusion function $i\colon P \to A^\star$ (given by $i(x) = x$). The induction assumptions on P essentially say that P carries an algebra structure $\mathsf{nil}\colon 1 \to P$, $\mathsf{cons}\colon A \times P \to P$, in such a way that the inclusion map $i\colon P \to A^\star$ is a map of algebras:

$$\begin{array}{ccc} 1 + (A \times P) & \xrightarrow{id + (id \times i)} & 1 + (A \times A^\star) \\ {\scriptstyle [\mathsf{nil}, \mathsf{cons}]}\downarrow & & \cong \downarrow {\scriptstyle [\mathsf{nil}, \mathsf{cons}]} \\ P & \xrightarrow[i]{} & A^\star \end{array}$$

In other words: P is a subalgebra of $A^\star$. By initiality we get a function $j\colon A^\star \to P$ as on the left below. But then $i \circ j = id$, by uniqueness.

$$\begin{array}{ccccc} & & id + (id \times id) & & \\ & \overset{\curvearrowright}{} & & \overset{\curvearrowright}{} & \\ 1 + (A \times A^\star) & \xrightarrow[id + (id \times j)]{} & 1 + (A \times P) & \xrightarrow[id + (id \times i)]{} & 1 + (A \times A^\star) \\ {\scriptstyle [\mathsf{nil}, \mathsf{cons}]}\downarrow \cong & & \downarrow {\scriptstyle [\mathsf{nil}, \mathsf{cons}]} & & \cong \downarrow {\scriptstyle [\mathsf{nil}, \mathsf{cons}]} \\ A^\star & \xrightarrow{\quad j \quad} & P & \xrightarrow{\quad i \quad} & A^\star \\ & \underset{\curvearrowright}{} & & \underset{\curvearrowright}{} & \\ & & id & & \end{array}$$

This means that $P = A^\star$, as we wished to derive.

4. The initiality property from Definition 5.3 allows us to define functions $f\colon U \to V$ out of an initial algebra (with carrier) U. Often one wishes to define functions $U \times D \to V$ involving an additional parameter ranging over a set D. A typical example is the addition function $\mathsf{plus}\colon \mathbb{N} \times \mathbb{N} \to \mathbb{N}$, defined by induction on (say) its first argument, with the second argument as parameter. One can handle such functions $U \times D \to V$ via Currying: they correspond to functions $U \to V^D$. And the latter can be defined via the initiality scheme. For example, we can define a Curryied addition function $\mathsf{plus}\colon \mathbb{N} \to \mathbb{N}^{\mathbb{N}}$ via initiality by putting an appropriate algebra structure $1 + \mathbb{N}^{\mathbb{N}} \to \mathbb{N}^{\mathbb{N}}$ on $\mathbb{N}$ (see Example 5.4):

$$\begin{array}{ccc} 1 + \mathbb{N} & \xrightarrow{\mathsf{plus}} & 1 + \mathbb{N}^{\mathbb{N}} \\ {\scriptstyle [0, S]}\downarrow \cong & & \downarrow {\scriptstyle [\lambda x.\, x,\ \lambda f.\, \lambda x.\, S(f(x))]} \\ \mathbb{N} & \xrightarrow[\mathsf{plus}]{} & \mathbb{N}^{\mathbb{N}} \end{array}$$

This says that

$$\mathsf{plus}(0) = \lambda x.\, x \qquad \text{and} \qquad \mathsf{plus}(n+1) = \lambda x.\, S(\mathsf{plus}(n)(x)).$$

Alternatively, one may formulate initiality "with parameters", see [15, 33], so that one can handle such functions $U \times D \to V$ directly.

6 Coalgebras and coinduction

In Section 4 we have seen that a "co"-product $+$ behaves like a product $\times$, except that the arrows point in opposite direction: one has coprojections $X \to X + Y \leftarrow Y$ instead of projections $X \leftarrow X \times Y \to Y$, and cotupleing instead of tupleing. One says that the coproduct $+$ is the dual of the product $\times$, because the associated arrows are reversed. Similarly, a "co"-algebra is the dual of an algebra.

Definition 6.1 For a functor T, a *coalgebra* (or a T-*coalgebra*) is a pair (U, c) consisting of a set U and a function $c\colon U \to T(U)$.

Like for algebras, we call the set U the *carrier* and the function c the *structure* or *operation* of the coalgebra (U, c). Because coalgebras often describe dynamical systems (of some sort), the carrier set U is also called the *state space*.

What, then, is the difference between an algebra $T(U) \to U$ and a coalgebra $U \to T(U)$? Essentially, it is the difference between construction and observation. An algebra consists of a carrier set U with a function $T(U) \to U$ going *into* this carrier U. It tells us how to construct elements in U. And a coalgebra consists of a carrier set U with a function $U \to T(U)$ in the opposite direction, going *out of* U. In this case we do not know how to form elements in U, but we only have operations acting on U, which may give us some information about U. In general, these coalgebraic operations do not tell us all there is to say about elements of U, so that we only have *limited access* to U. Coalgebras—like algebras—can be seen as models of a signature of operations—not of constructor operations, but of destructor/observer operations.

Consider for example the functor $T(X) = A \times X$, where A is a fixed set. A coalgebra $U \to T(U)$ consists of two functions $U \to A$ and $U \to U$, which we earlier called $\mathsf{value}\colon U \to A$ and $\mathsf{next}\colon U \to U$. With these operations we can do two things, given an element $u \in U$:

1. produce an element in A, namely $\mathsf{value}(u)$;
2. produce a next element in U, namely $\mathsf{next}(u)$.

Now we can repeat 1. and 2. and form another element in A, namely $\mathsf{value}(\mathsf{next}(u))$. By proceeding in this way we can get for each element $u \in U$ an infinite sequence $(a_1, a_2, \ldots) \in A^{\mathbb{N}}$ of elements of $a_i = \mathsf{value}(\mathsf{next}^{(n)}(u)) \in A$. This sequence of elements that u gives rise to is what we can *observe* about u. Two elements $u_1, u_2 \in U$ may well give rise to the same sequence of elements of A, without actually being equal as elements of U. In such a case one calls u_1 and u_2 observationally indistinguishable, or bisimilar.

Here is another example. Let the functor $T(X) = 1 + A \times X$ have a coalgebra $\mathsf{pn}\colon U \to 1 + A \times U$, where '$\mathsf{pn}$' stands for 'possible next'. If we have an element $u \in U$, then we can see the following.

1. Either $\mathsf{pn}(u) = \kappa(*) \in 1 + A \times U$ is in the left component of $+$. If this happens, then our experiment stops, since there is no state (element of U) left with which to continue.
2. Or $\mathsf{pn}(u) = \kappa'(a, u) \in 1 + A \times U$ is in the right $+$-component. This gives us an element $a \in A$ and a next element $u' \in U$ of the carrier, with which we can proceed.

Repeating this we can observe for an element $u \in U$ either a finite sequence $(a_1, a_2, \ldots, a_n) \in A^\star$ of, or an infinite sequence $(a_1, a_2, \ldots) \in A^{\mathbb{N}}$. The observable outcomes are elements of the set $A^\infty = A^\star + A^{\mathbb{N}}$ of finite and infinite lists of A's.

These observations will turn out to be elements of the final coalgebra of the functors involved, see Example 6.3 and 6.5 below. But in order to formulate this notion of finality for coalgebras we first need to know what a "homomorphism of coalgebras" is. It is, like in algebra, a function between the underlying sets which commutes with the operations. For example, let $T(X) =$

$A \times X$ be the "infinite list" functor as used above, with coalgebras $\langle h_1, t_1 \rangle : U_1 \to A \times U_1$ and $\langle h_2, t_2 \rangle : U_2 \to A \times U_2$. A homomorphism of coalgebras from the first to the second consists of a function $f : U_1 \to U_2$ between the carrier sets (state spaces) with $h_2 \circ f = h_1$ and $t_2 \circ f = f \circ t_1$ in:

$$\begin{array}{ccc} U_1 & \xrightarrow{f} & U_2 \\ {\scriptstyle h_1}\downarrow & & \downarrow{\scriptstyle h_2} \\ A & = & A \end{array} \qquad \text{and} \qquad \begin{array}{ccc} U_1 & \xrightarrow{f} & U_2 \\ {\scriptstyle t_1}\downarrow & & \downarrow{\scriptstyle t_2} \\ U_1 & \xrightarrow[f]{} & U_2 \end{array}$$

These two diagrams can be combined into a single one:

$$\begin{array}{ccc} U_1 & \xrightarrow{f} & U_2 \\ {\scriptstyle \langle h_1, t_1 \rangle}\downarrow & & \downarrow{\scriptstyle \langle h_2, t_2 \rangle} \\ A \times U_1 & \xrightarrow[id \times f]{} & A \times U_2 \end{array} \qquad \text{i.e. into} \qquad \begin{array}{ccc} U_1 & \xrightarrow{f} & U_2 \\ {\scriptstyle \langle h_1, t_1 \rangle}\downarrow & & \downarrow{\scriptstyle \langle h_2, t_2 \rangle} \\ T(U_1) & \xrightarrow[T(f)]{} & T(U_2) \end{array}$$

Definition 6.2 Let T be a functor.

(i) A *homomorphism of coalgebras* (or, *map of coalgebras*, or *coalgebra map*) from a T-coalgebra $U_1 \xrightarrow{c_1} T(U_1)$ to another T-coalgebra $U_2 \xrightarrow{c_2} T(U_2)$ consists of a function $f : U_1 \to U_2$ between the carrier sets which commutes with the operations: $c_2 \circ f = T(f) \circ c_1$ as expressed by the following diagram.

$$\begin{array}{ccc} U_1 & \xrightarrow{f} & U_2 \\ {\scriptstyle c_1}\downarrow & & \downarrow{\scriptstyle c_2} \\ T(U_1) & \xrightarrow[T(f)]{} & T(U_2) \end{array}$$

(ii) A *final coalgebra* $d : W \to T(W)$ is a coalgebra such that for every coalgebra $c : U \to T(U)$ there is a unique map of coalgebras $(U, C) \to (W, d)$.

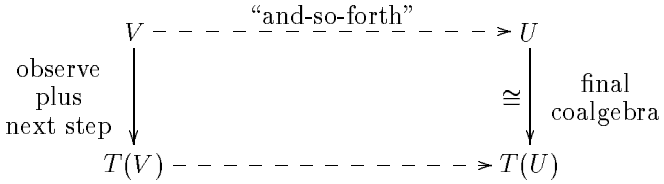
Notice that where the initiality property for algebras allows us to define functions going *out of* an initial algebra, the finality property for coalgebras gives us means to define functions *into* a final coalgebra. Earlier we have emphasised that what is typical in a coalgebraic setting is that there are no operations for constructing elements of a state space (of a coalgebra), and that state spaces should therefore be seen as black boxes. However, if we know that a certain coalgebra is final, then we can actually form elements in its state space by this finality principle. The next example contains some illustrations. Besides a means for constructing elements, finality also allows us to define various operations on final coalgebras, as will be shown in a series of examples below. (In fact, in this way one can put certain algebraic structure on top of a coalgebra, see [67] for a systematic study in the context of process algebras.)

Now that we have seen the definitions of initiality (for algebras, see Definition 5.3) and finality (for coalgebras) we are in a position to see the formal similarities. At an informal level we can explain these similarities as follows. A typical initiality diagram may be drawn as:

$$\begin{array}{ccc} T(U) & \dashrightarrow & T(V) \\ \text{initial algebra}\downarrow{\scriptstyle \cong} & & \downarrow \text{base step plus next step} \\ U & \dashrightarrow & V \end{array}$$

"and-so-forth"

The map "and-so-forth" that is defined in this diagram applies the "next step" operations repeatedly to the "base step". The pattern in a finality diagram is similar:



In this case the "and-so-forth" map captures the observations that arise by repeatedly applying the "next step" operation. This captures the observable behaviour.

The technique for defining a function $f\colon V \to U$ by finality is thus: describe the direct observations together with the single next steps of f as a coalgebra structure on V. The function f then arises by repetition. Hence a coinductive definition of f does not determine f "at once", but "step-by-step". In the next section we shall describe proof techniques using bisimulations, which fully exploit this step-by-step character of coinductive definitions.

But first we identify a simply coalgebra concretely, and show how we can use finality.

Example 6.3 For a fixed set A, consider the functor $T(X) = A \times X$. We claim that the final coalgebra of this functor is the set $A^{\mathbb{N}}$ of infinite lists of elements from A, with coalgebra structure

$$\langle \mathsf{head}, \mathsf{tail} \rangle \colon A^{\mathbb{N}} \longrightarrow A \times A^{\mathbb{N}}$$

given by

$$\mathsf{head}(\alpha) = \alpha(0) \qquad \text{and} \qquad \mathsf{tail}(\alpha) = \lambda x.\, \alpha(x+1).$$

Hence head takes the first element of an infinite sequence $(\alpha(0), \alpha(1), \alpha(2), \ldots)$ of elements of A, and tail takes the remaining list. We notice that the pair of functions $\langle \mathsf{head}, \mathsf{tail} \rangle \colon A^{\mathbb{N}} \to A \times A^{\mathbb{N}}$ is an isomorphism.

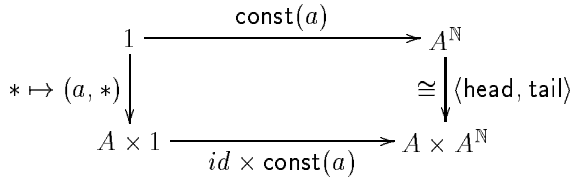
We claim that for an arbitrary coalgebra $\langle \mathsf{value}, \mathsf{next} \rangle \colon U \to A \times U$ there is a unique homomorphism of coalgebras $f\colon U \to A^{\mathbb{N}}$; it is given for $u \in U$ and $n \in \mathbb{N}$ by

$$f(u)(n) = \mathsf{value}\left(\mathsf{next}^{(n)}(u)\right).$$

Then indeed, $\mathsf{head} \circ f = \mathsf{value}$ and $\mathsf{tail} \circ f = f \circ \mathsf{next}$, making f a map of coalgebras. And f is unique in satisfying these two equations, as can be checked easily.

Earlier in this section we saw that what we can observe about an element $u \in U$ is an infinite list of elements of A arising as $\mathsf{value}(u)$, $\mathsf{value}(\mathsf{next}(u))$, $\mathsf{value}(\mathsf{next}(\mathsf{next}(u)))$, ... Now we see that this observable behaviour of u is precisely the outcome $f(u) \in A^{\mathbb{N}}$ at u of the unique map f to the final coalgebra. Hence the elements of the final coalgebra give the observable behaviour. This is typical for final coalgebras.

Once we know that $A^{\mathbb{N}}$ is a final coalgebra—or, more precisely, carries a final coalgebra structure—we can use this finality to define functions into $A^{\mathbb{N}}$. Let us start with a simple example, which involves defining the constant sequence $\mathsf{const}(a) = (a, a, a, \ldots) \in A^{\mathbb{N}}$ by coinduction (for some element $a \in A$). We shall define this constant as a function $\mathsf{const}(a)\colon 1 \to A^{\mathbb{N}}$, where $1 = \{*\}$ is a singleton set. Following the above explanation, we have to produce a coalgebra structure $1 \to T(1) = A \times 1$ on 1, in such a way that $\mathsf{const}(a)$ arises by repetition. In this case the only thing we want to observe is the element $a \in A$ itself, and so we simply define as coalgebra structure $1 \to A \times 1$ the function $* \mapsto (a, *)$. Indeed, $\mathsf{const}(a)$ arises in the following finality diagram.



It expresses that $\mathsf{head}(\mathsf{const}(a)) = a$ and $\mathsf{tail}(\mathsf{const}(a)) = \mathsf{const}(a)$.

We consider another example, for the special case where $A = \mathbb{N}$. We now wish to define (coinductively) the function $\mathsf{from}\colon \mathbb{N} \to \mathbb{N}^{\mathbb{N}}$ which maps a natural number $n \in \mathbb{N}$ to the sequence $(n, n+1, n+2, n+3, \ldots) \in \mathbb{N}^{\mathbb{N}}$. This involves defining a coalgebra structure $\mathbb{N} \to \mathbb{N} \times \mathbb{N}$ on the domain $\mathbb{N}$ of the function from that we are trying to define. The direct observation that we can make about a "state" $n \in \mathbb{N}$ is n itself, and the next state is then $n+1$ (in which we can directly observe $n+1$). Repetition then leads to $\mathsf{from}(n)$. Thus we define the function from in the following diagram.

$$\begin{array}{ccc} \mathbb{N} & \xrightarrow{\quad\mathsf{from}\quad} & \mathbb{N}^{\mathbb{N}} \\ {\scriptstyle \lambda n.\,(n, n+1)}\downarrow & & \cong\downarrow{\scriptstyle \langle \mathsf{head}, \mathsf{tail}\rangle} \\ \mathbb{N} \times \mathbb{N} & \xrightarrow[id \times \mathsf{from}]{} & \mathbb{N} \times \mathbb{N}^{\mathbb{N}} \end{array}$$

It is then determined by the equations $\mathsf{head}(\mathsf{from}(n)) = n$ and $\mathsf{tail}(\mathsf{from}(n)) = \mathsf{from}(n+1)$.

We are now in a position to provide the formal background for the examples of coinductive definitions and proofs in Section 3. For instance, the function $\mathsf{merge}\colon A^{\mathbb{N}} \times A^{\mathbb{N}} \to A^{\mathbb{N}}$ which merges two infinite lists into a single one arises as unique function to the final coalgebra $A^{\mathbb{N}}$ in:

$$\begin{array}{ccc} A^{\mathbb{N}} \times A^{\mathbb{N}} & \xrightarrow{\quad\mathsf{merge}\quad} & A^{\mathbb{N}} \\ {\scriptstyle \lambda(\alpha,\beta).\,(\mathsf{head}(\alpha),(\beta,\mathsf{tail}(\alpha)))}\downarrow & & \cong\downarrow{\scriptstyle \langle \mathsf{head}, \mathsf{tail}\rangle} \\ A \times (A^{\mathbb{N}} \times A^{\mathbb{N}}) & \xrightarrow[id \times \mathsf{merge}]{} & A \times A^{\mathbb{N}} \end{array}$$

Notice that the coalgebra structure on the left that we put on the domain $A^{\mathbb{N}} \times A^{\mathbb{N}}$ of merge corresponds to the defining "coinduction" clauses for merge, as used in Section 3. It expresses the direct observation after a merge, together with the next state (about which we make a next direct observation).

It follows from the commutativity of the above diagram that

$$\mathsf{head}(\mathsf{merge}(\alpha,\beta)) = \mathsf{head}(\alpha) \qquad \text{and} \qquad \mathsf{tail}(\mathsf{merge}(\alpha,\beta)) = \mathsf{merge}(\beta, \mathsf{tail}(\alpha)).$$

The function $\mathsf{odd} : A^{\mathbb{N}} \to A^{\mathbb{N}}$ can similarly be defined coinductively, that is, by finality of $A^{\mathbb{N}}$, as follows:

$$\begin{array}{ccc} A^{\mathbb{N}} & \xrightarrow{\quad\mathsf{odd}\quad} & A^{\mathbb{N}} \\ {\scriptstyle \lambda\alpha.\,(\mathsf{head}(\alpha),\mathsf{tail}(\mathsf{tail}(\alpha)))}\downarrow & & \cong\downarrow{\scriptstyle \langle \mathsf{head}, \mathsf{tail}\rangle} \\ A \times A^{\mathbb{N}} & \xrightarrow[id \times \mathsf{odd}]{} & A \times A^{\mathbb{N}} \end{array}$$

The coalgebra structure on $A^{\mathbb{N}}$ on the left gives by finality rise to a unique coalgebra homomorphism, called odd. By the commutatitivity of the diagram, it satisfies:

$$\mathsf{head}(\mathsf{odd}(\alpha)) = \mathsf{head}(\alpha) \qquad \text{and} \qquad \mathsf{tail}(\mathsf{odd}(\alpha)) = \mathsf{odd}(\mathsf{tail}(\mathsf{tail}(\alpha))).$$

As before, we define

$$\mathsf{even}(\alpha) = \mathsf{odd}(\mathsf{tail}(\alpha)).$$

Next we prove for all α in $A^{\mathbb{N}}$: $\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha)) = \alpha$, by showing that $\mathsf{merge} \circ \langle \mathsf{odd}, \mathsf{even}\rangle$ is a homomorphism of coalgebras from $(A^{\mathbb{N}}, \langle \mathsf{head}, \mathsf{tail}\rangle)$ to $(A^{\mathbb{N}}, \langle \mathsf{head}, \mathsf{tail}\rangle)$. The required equality then follows by uniqueness, because the identity function $id\colon A^{\mathbb{N}} \to A^{\mathbb{N}}$ is (trivially) a homomorphism $(A^{\mathbb{N}}, \langle \mathsf{head}, \mathsf{tail}\rangle) \to (A^{\mathbb{N}}, \langle \mathsf{head}, \mathsf{tail}\rangle)$ as well. Thus, all we have to prove is that we have a homomorphism, i.e. that

$$\langle \mathsf{head}, \mathsf{tail}\rangle \circ (\mathsf{merge} \circ \langle \mathsf{odd}, \mathsf{even}\rangle) = (id \times (\mathsf{merge} \circ \langle \mathsf{odd}, \mathsf{even}\rangle)) \circ \langle \mathsf{head}, \mathsf{tail}\rangle.$$

This follows from the following two computations.

$$\begin{aligned} \mathsf{head}(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha))) &= \mathsf{head}(\mathsf{odd}(\alpha)) \\ &= \mathsf{head}(\alpha). \end{aligned}$$

And:

$$\begin{aligned} \mathsf{tail}(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha))) &= \mathsf{merge}(\mathsf{even}(\alpha), \mathsf{tail}(\mathsf{odd}(\alpha))) \\ &= \mathsf{merge}(\mathsf{even}(\alpha), \mathsf{odd}(\mathsf{tail}(\mathsf{tail}(\alpha)))) \\ &= \mathsf{merge}(\mathsf{odd}(\mathsf{tail}(\alpha)), \mathsf{even}(\mathsf{tail}(\alpha))) \\ &= (\mathsf{merge} \circ \langle \mathsf{odd}, \mathsf{even} \rangle)(\mathsf{tail}(\alpha)). \end{aligned}$$

In Section 7, an alternative method for proving facts such as the one above, will be introduced, which is based on the notion of bisimulation.

Clearly, there are formal similarities between algebra maps and coalgebra maps. We leave it to the reader to check that coalgebra maps can be composed as functions, and that the identity function on the carrier of a coalgebra is a map of coalgebras. There is also the following result, which is dual—including its proof—to Lemma 5.5.

Lemma 6.4 *(i) Final coalgebras, if they exist, are uniquely determined (up-to-isomorphism).*
(ii) A final coalgebra $W \to T(W)$ is a fixed point $W \xrightarrow{\cong} T(W)$ of the functor T. □

Final coalgebras are generalizations of greatest fixed points of monotone functions. As for initial algebras, the existence of final coalgebras is more important than their actual (internal) structure. Their use is determined entirely by their finality property, and not by their structure. Often, the existence of a final coalgebra follows from general properties of the relevant functor (and of the underlying category), see e.g. [43, 66].

The unique existence of a map of coalgebras into a final coalgebra has two aspects: existence, which gives us a principle of definition by coinduction, and uniqueness, which gives us a principle of proof by coinduction. This will be further illustrated in a series of examples, which will occupy the remainder of this section.

Example 6.5 It is not hard to show that the final coalgebra of the functor $T(X) = 1 + (A \times X)$ has as carrier the set $A^\infty = A^\star + A^{\mathbb{N}}$ of finite and infinite lists of A's. The associated "possible next" coalgebra structure

$$\mathsf{pn}\colon A^\infty \longrightarrow 1 + A \times A^\infty \qquad \text{is} \qquad \alpha \mapsto \begin{cases} \kappa(*) & \text{if } \alpha = () \\ \kappa'(a, \alpha') & \text{if } \alpha = a \cdot \alpha' \end{cases}$$

is final: for an arbitrary coalgebra $g\colon U \to 1 + (A \times U)$ of the functor T there is a unique homomorphism of coalgebras $f\colon U \to A^\infty$. Earlier in this section we identified such lists in A^∞ as the observable behaviour for machines whose signature of operations is described by T.

We give some examples of coinductive definitions for such finite and infinite lists. First an easy one, describing an empty list $\mathsf{nil}\colon 1 \to A^\infty$ as the unique coalgebra homomorphisms in the following situation.

$$\begin{array}{ccc} 1 & \xrightarrow{\quad\mathsf{nil}\quad} & A^\infty \\ {\scriptstyle\kappa}\downarrow & & \cong\downarrow{\scriptstyle\mathsf{pn}} \\ 1 + (A \times 1) & \xrightarrow[id + (id \times \mathsf{nil})]{} & 1 + (A \times A^\infty) \end{array}$$

This determines nil as $\mathsf{pn}^{-1} \circ \kappa$. We define a prefix operation $\mathsf{cons}\colon A \times A^\infty \to A^\infty$ as $\mathsf{pn}^{-1} \circ \kappa'$.

We can coinductively define a list inclusion function $\mathsf{list\_incl}\colon A^\star \to A^\infty$ via the coalgebra structure $\mathsf{list\_incl\_struct}\colon A^\star \to 1 + (A \times A^\star)$ given by

$$\alpha \mapsto \begin{cases} \kappa(*) & \text{if } \alpha = \mathsf{nil} \\ \kappa'(a, \beta) & \text{if } \alpha = \mathsf{cons}(a, \beta) \end{cases}$$

We leave it to the reader to (coinductively) define an infinite list inclusion $A^{\mathbb{N}} \to A^\infty$.

A next, more serious example, involves the concatenation function $\mathsf{conc}\colon A^\infty \times A^\infty \to A^\infty$ which yields for two lists $x, y \in A^\infty$ a new list $\mathsf{cons}(x,y) \in A^\infty$ which contains the elements of x followed by the elements of y. Coinductively one defines $\mathsf{conc}(x,y)$ by laying down what the possible observations are on this new list $\mathsf{conc}(x,y)$. Concretely, this means that we should define what $\mathsf{pn}(\mathsf{conc}(x,y))$ is. The intuition we have of concatenation tells us that the possible next $\mathsf{pn}(\mathsf{conc}(x,y))$ is the possible next $\mathsf{pn}(x)$ of x if x is not the empty list (i.e. if $\mathsf{pn}(x) \neq \kappa(*) \in 1$), and the possible next $\mathsf{pn}(y)$ of y otherwise. This is captured in the coalgebra structure $\mathsf{conc\_struct}\colon A^\infty \times A^\infty \to 1 + (A \times (A^\infty \times A^\infty))$ given by:

$$(\alpha,\beta) \mapsto \begin{cases} \kappa(*) & \text{if } \mathsf{pn}(\alpha) = \mathsf{pn}(\beta) = \kappa(*) \\ \kappa'(a,(\alpha',\beta)) & \text{if } \mathsf{pn}(\alpha) = \kappa'(a,\alpha') \\ \kappa'(b,(\alpha,\beta')) & \text{if } \mathsf{pn}(\alpha) = \kappa(*) \text{ and } \mathsf{pn}(\beta) = \kappa'(b,\beta'). \end{cases}$$

The concatenation function $\mathsf{conc}\colon A^\infty \times A^\infty \to A^\infty$ that we wished to define arises as unique coalgebra map resulting from $\mathsf{conc\_struct}$.

The interested reader may wish to prove (by uniqueness!) that:

$$\mathsf{conc}(x,\mathsf{nil}) = x = \mathsf{conc}(\mathsf{nil},x)$$
$$\mathsf{conc}(\mathsf{conc}(x,y),z) = \mathsf{conc}(x,\mathsf{conc}(y,z)).$$

One may also wish to prove that $\mathsf{conc}(\mathsf{cons}(a,x),y)) = \mathsf{cons}(a,\mathsf{conc}(x,y))$. The easiest way is to show that applying pn on both sides yields the same result. Then we are done, since pn is an isomorphism.

Example 6.6 Consider the functor $T(X) = 1 + X$ from Example 5.4. Remember that its initial algebra is given by the set $\mathbb{N} = \{0,1,2,\ldots,\}$ of natural numbers with cotuple of zero and successor functions as algebra structure $[0,S]\colon 1 + \mathbb{N} \xrightarrow{\cong} \mathbb{N}$.

The final coalgebra $\overline{\mathbb{N}} \xrightarrow{\cong} 1 + \overline{\mathbb{N}}$ of T is the set

$$\overline{\mathbb{N}} = \{0,1,2,\ldots,\} \cup \{\infty\}$$

of natural numbers augmented with an extra element ∞. The final coalgebra structure $\overline{\mathbb{N}} \to 1 + \overline{\mathbb{N}}$ is best called a predecessor pred because it sends

$$0 \mapsto \kappa(*), \qquad n+1 \mapsto \kappa'(n), \qquad \infty \mapsto \kappa'(\infty)$$

where we have written the coprojections κ, κ' explicitly in order to emphasise the $+$-component to which $\mathsf{pred}(x) \in 1 + \overline{\mathbb{N}}$ belongs. This final coalgebra may be obtained by taking as the constant set A a singleton set 1 for the functor $X \mapsto 1 + (A \times X)$ in the previous example. And indeed, the set $1^\infty = 1^\star + 1^{\mathbb{N}}$ is isomorphic to $\overline{\mathbb{N}}$. The "possible next" operations $\mathsf{pn}\colon 1^\infty \to 1 + (1 \times 1^\infty)$ is then indeed the predecessor.

The defining property of this final coalgebra $\mathsf{pred}\colon \overline{\mathbb{N}} \to 1 + \overline{\mathbb{N}}$ says that for every set U with a function $f\colon U \to 1 + U$ there is a unique function $g\colon U \to \overline{\mathbb{N}}$ in the following diagram.

$$\begin{array}{ccc} U & \overset{g}{\dashrightarrow} & \overline{\mathbb{N}} \\ {\scriptstyle f}\downarrow & & \cong\downarrow{\scriptstyle \mathsf{pred}} \\ 1 + U & \underset{id+g}{\dashrightarrow} & 1 + \overline{\mathbb{N}} \end{array}$$

This says that g is the unique function satisfying

$$\mathsf{pred}(g(x)) = \begin{cases} \kappa(*) & \text{if } f(x) = \kappa(*) \\ \kappa'(g(x')) & \text{if } f(x) = \kappa'(x'). \end{cases}$$

This function g gives us the behaviour that one can observe about systems with one button $X \to 1 + X$, as mentioned in the first (coalgebra) example in Section 2.

Consider now the function $f\colon \overline{\mathbb{N}} \times \overline{\mathbb{N}} \to 1 + (\overline{\mathbb{N}} \times \overline{\mathbb{N}})$ defined by

$$f(x,y) = \begin{cases} \kappa(*) & \text{if } \mathsf{pred}(x) = \mathsf{pred}(y) = \kappa(*) \\ \kappa'(\langle x', y\rangle) & \text{if } \mathsf{pred}(x) = \kappa'(x') \\ \kappa'(\langle x, y'\rangle) & \text{if } \mathsf{pred}(x) = \kappa(*),\ \mathsf{pred}(y) = \kappa'(y'). \end{cases}$$

This f puts a coalgebra structure on $\overline{\mathbb{N}} \times \overline{\mathbb{N}}$, for the functor $X \mapsto 1 + X$ that we are considering. Hence it gives rise to a unique coalgebra homomorphism $\oplus\colon \overline{\mathbb{N}} \times \overline{\mathbb{N}} \to \overline{\mathbb{N}}$ in the following situation.

$$\begin{array}{ccc} \overline{\mathbb{N}} \times \overline{\mathbb{N}} & \xrightarrow{\ \oplus\ } & \overline{\mathbb{N}} \\ {\scriptstyle f}\downarrow & & \cong\downarrow{\scriptstyle \mathsf{pred}} \\ 1 + (\overline{\mathbb{N}} \times \overline{\mathbb{N}}) & \xrightarrow[id + \oplus]{} & 1 + \overline{\mathbb{N}} \end{array}$$

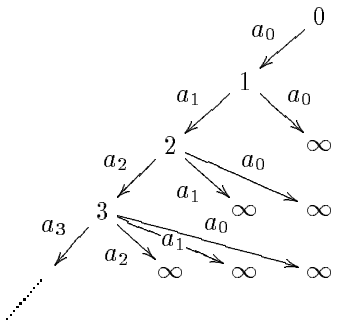
Hence $\oplus$ is the unique function $\overline{\mathbb{N}} \times \overline{\mathbb{N}} \to \overline{\mathbb{N}}$ with

$$\mathsf{pred}(x \oplus y) = \begin{cases} \kappa(*) & \text{if } \mathsf{pred}(x) = \kappa(*) = \mathsf{pred}(y) \\ \kappa'(x \oplus y') & \text{if } \mathsf{pred}(x) = \kappa(*), \mathsf{pred}(y) = \kappa'(y') \\ \kappa'(x' \oplus y) & \text{if } \mathsf{pred}(x) = \kappa'(x'). \end{cases}$$

It is not hard to see that $n \oplus m = n + m$ for $n, m \in \mathbb{N}$ and $n \oplus \infty = \infty = \infty \oplus n$, so that $\oplus$ behaves like addition on the "extended" natural numbers in $\overline{\mathbb{N}}$. One easily verifies that this addition function $\oplus\colon \overline{\mathbb{N}} \times \overline{\mathbb{N}} \to \overline{\mathbb{N}}$ is the special case (for $A = 1$) of the concatenation function $\mathsf{conc}\colon A^\infty \times A^\infty \to A^\infty$ that we introduced in the previous example. This special case distinguishes itself in an important aspect: it can be shown that concatenation (or addition) $\oplus\colon \overline{\mathbb{N}} \times \overline{\mathbb{N}} \to \overline{\mathbb{N}}$ on the extended natural numbers is commutative—e.g. by uniqueness, or by bisimulation (see [65] for details)—whereas concatenation $\mathsf{conc}\colon A^\infty \times A^\infty \to A^\infty$ in general is not commutative. If A has more than two elements, then $\mathsf{conc}(x,y) \neq \mathsf{conc}(y,x)$, because they give rise to different observations, e.g. for both x, y singleton sequence containing different elements.

Example 6.7 We consider the functor $T(X) = \mathsf{list}(A \times X)$, for a constant set A. On a function $f\colon X \to Y$ it yields a function $\mathsf{list}(A \times X) \to \mathsf{list}(A \times Y)$ sending $\langle (a_1, x_1), \ldots, (a_n, x_n)\rangle$ to $\langle (a_1, f(x_1)), \ldots, (a_n, f(x_n))\rangle$. A coalgebra $X \to \mathsf{list}(A \times X)$ of this functor is a function which maps a state $x \in X$ to a finite list of pairs (a_i, x_i) consisting of a label $a_i \in A$ and a successor state $x_i \in X$. Since this passage to next states can be repeated, such a coalgebra describes a tree of possibly infinite depth, with finitely many ordered branches, each provided with a label from A.

For example, the following tree with infinite set $(a_n)_{n \in \mathbb{N}}$ of labels:



can be described as a coalgebra $t\colon \overline{\mathbb{N}} \to \mathsf{list}(A \times \overline{\mathbb{N}})$ on the state space $\overline{\mathbb{N}} = \mathbb{N} \cup \{\infty\}$ as:

$$t(\infty) = \langle\rangle = \mathsf{nil} \qquad \text{and} \qquad t(n) = \langle (a_n, n+1), (a_{n-1}, \infty), \ldots, (a_0, \infty)\rangle.$$

Our aim in this example is to define for an arbitrary T-coalgebra $t\colon X \to \mathsf{list}(A \times X)$ two classical traversal "algorithms" $\mathsf{bf}(-)$ and $\mathsf{df}(-)$ for breadth-first and depth-first, yielding functions $\mathsf{bf}(t)\colon X \to A^\infty$ and $\mathsf{df}(t)\colon X \to A^\infty$ which describe the elements occurring in the tree t as element in a possibly infinite list, either in breadth-first or in depth-first order. These functions $\mathsf{bf}(t)$ and $\mathsf{df}(t)$ are defined using the finality of $A^\infty = A^\star + A^{\mathbb{N}}$, which is the (carrier of the) final coalgebra of the functor $X \mapsto 1 + (A \times X)$ (see Example 6.5). This is done in two steps: we first put two coalgebra structures $\mathsf{bfs}(t)\colon \mathsf{list}(A \times X) \to 1 + (A \times \mathsf{list}(A \times X))$ and $\mathsf{dfs}(t)\colon \mathsf{list}(A \times X) \to 1 + (A \times \mathsf{list}(A \times X))$ on $\mathsf{list}(A \times X)$. By finality, these will give rise to two coalgebra homomorphisms $\mathsf{bfh}(t)\colon \mathsf{list}(A \times X) \to A^\infty$ and $\mathsf{dfh}(t)\colon \mathsf{list}(A \times X) \to A^\infty$. Then we define:

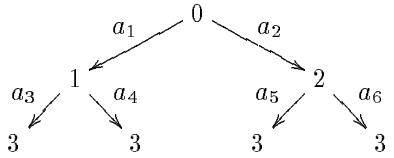
$$\mathsf{bf}(t) = \mathsf{bfh}(t) \circ t\colon X \longrightarrow A^\infty \qquad \text{and} \qquad \mathsf{df}(t) = \mathsf{dfh}(t) \circ t\colon X \longrightarrow A^\infty.$$

So the only thing we should do is define two coalgebra structures $\mathsf{bfs}(t)$ and $\mathsf{dfs}(t)$ on $\mathsf{list}(A \times X)$, as functions $\mathsf{list}(A \times X) \to 1 + (A \times \mathsf{list}(A \times X))$. We take:

$$\ell \overset{\mathsf{bfs}(t)}{\longmapsto} \begin{cases} * & \text{if } \ell = \mathsf{nil} \\ (a, \ell' \cdot t(x)) & \text{if } \ell = \mathsf{cons}((a,x),\ell') \end{cases} \qquad \ell \overset{\mathsf{dfs}(t)}{\longmapsto} \begin{cases} * & \text{if } \ell = \mathsf{nil} \\ (a, t(x) \cdot \ell') & \text{if } \ell = \mathsf{cons}((a,x),\ell') \end{cases}$$

where $\cdot$ is concatenation on $\mathsf{list}(A \times X)$.

We illustrate the resulting functions $\mathsf{bf}(t), \mathsf{df}(t)\colon X \to A^\infty$ in an example. Consider therefore the tree



on the state space $X = \{0,1,2,3\}$ with coalgebra structure $t\colon X \to \mathsf{list}(A \times X)$ given as

$$t(0) = \langle (a_1,1),(a_2,2)\rangle, \quad t(1) = \langle (a_3,3),(a_4,3)\rangle, \quad t(2) = \langle (a_5,3),(a_6,3)\rangle, \quad t(3) = \langle\rangle.$$

Then we can compute the two resulting breadth-first and depth-first traversals of the tree t as:

$$\begin{aligned}
\mathsf{bf}(t)(0) &= \mathsf{bfh}(t)(t(0)) \\
&= \mathsf{bfh}(t)(\langle (a_1,1),(a_2,2)\rangle) \\
&= a_1 \cdot \mathsf{bfh}(t)((a_2,2),(a_3,3),(a_4,3)\rangle) \\
&= a_1 \cdot a_2 \cdot \mathsf{bfh}(t)(\langle (a_3,3),(a_4,3),(a_5,3),(a_6,3)\rangle) \\
&= a_1 \cdot a_2 \cdot a_3 \cdot a_4 \cdot a_5 \cdot a_6. \\
\mathsf{df}(t)(0) &= \mathsf{dfh}(t)(\langle (a_1,1),(a_2,2)\rangle) \\
&= a_1 \cdot \mathsf{dfh}(t)((a_3,3),(a_4,3),(a_2,2)\rangle) \\
&= a_1 \cdot a_3 \cdot a_4 \cdot \mathsf{dfh}(t)(\langle (a_2,2)\rangle) \\
&= a_1 \cdot a_3 \cdot a_4 \cdot a_2 \cdot \mathsf{dfh}(t)(\langle (a_5,3),(a_6,3)\rangle) \\
&= a_1 \cdot a_3 \cdot a_4 \cdot a_2 \cdot a_5 \cdot a_6.
\end{aligned}$$

These calculations rely on the maps $\mathsf{bfh}(t)$ and $\mathsf{dfh}(t)$ being homomorphisms of coalgebras (of the functor $X \mapsto 1 + (A \times X)$). Defining "algorithms" like these by initiality or finality thus yields certain canonical equations, which can be used for reasoning, e.g. in correctness proofs. This has been developed into a programming methodology by Bird and Meertens, see the recent reference [12] for more information (mostly involving initiality) and further references. In this context, finality was first used in [47].

Actual programming purely on the basis of initiality and finality can be done in the language CHARITY, see [15] and the references mentioned there.

7 Proofs by coinduction and bisimulation

In this section, we shall give an alternative formulation for one of the earlier proofs by coinduction. The new proof does not (directly) exploit (the uniqueness aspect of) finality, but makes use of the notion of *bisimulation*. We also present one new examples and then formulate the general case, allowing us to proof equalities on final coalgebras via bisimulations.

We recall from Example 6.3 that the final coalgebra of the functor $T(X) = A \times X$ is the set of infinite lists $A^{\mathbb{N}}$ of elements of A with coalgebra structure $\langle \mathsf{head}, \mathsf{tail} \rangle$. A *bisimulation* on this carrier $A^{\mathbb{N}}$ is a relation R on $A^{\mathbb{N}}$ satisfying

$$R(\alpha, \beta) \Rightarrow \begin{cases} \mathsf{head}(\alpha) = \mathsf{head}(\beta), & \text{and} \\ R(\mathsf{tail}(\alpha), \mathsf{tail}(\beta)). \end{cases}$$

Now $A^{\mathbb{N}}$ satisfies the following *coinductive proof principle*, or cpp for short: For all α and β in $A^{\mathbb{N}}$,

$$\text{if } R(\alpha, \beta), \text{ for some bisimulation } R \text{ on } A^{\mathbb{N}}, \text{ then } \alpha = \beta. \tag{cpp}$$

Before we give a proof of the principle, which will be based on the finality of $A^{\mathbb{N}}$, we illustrate its use by proving, once again, for all α in $A^{\mathbb{N}}$,

$$\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha)) = \alpha.$$

To this end, define the following relation on $A^{\mathbb{N}}$:

$$R = \{(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha)),\ \alpha) \mid \alpha \in A^{\mathbb{N}}\}.$$

In order to prove the above equality it is, by the coinductive proof principle (cpp), sufficient to show that R is a bisimulation. This follows by proving the two requirements mentioned above. First, for each pair $(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha)), \alpha)$ in R we have equal head's:

$$\begin{aligned} \mathsf{head}(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha))) &= \mathsf{head}(\mathsf{odd}(\alpha)) \\ &= \mathsf{head}(\alpha). \end{aligned}$$

And secondly, if we have a pair $(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha)), \alpha)$ in R, then applying tail on both sides yields a new pair in R, since we can rewrite, using that $\mathsf{even} = \mathsf{odd} \circ \mathsf{tail}$,

$$\begin{aligned} \mathsf{tail}(\mathsf{merge}(\mathsf{odd}(\alpha), \mathsf{even}(\alpha))) &= \mathsf{merge}(\mathsf{even}(\alpha), \mathsf{tail}(\mathsf{odd}(\alpha))) \\ &= \mathsf{merge}(\mathsf{odd}(\mathsf{tail}(\alpha)), \mathsf{odd}(\mathsf{tail}(\mathsf{tail}(\alpha)))) \\ &= \mathsf{merge}(\mathsf{odd}(\mathsf{tail}(\alpha)), \mathsf{even}(\mathsf{tail}(\alpha)). \end{aligned}$$

For a proof of the cpp, let R be any bisimulation on $A^{\mathbb{N}}$. If we consider R as a set (of pairs), then it can be supplied with a $A \times (-)$-coalgebra structure by defining a function

$$\gamma \colon R \longrightarrow A \times R \qquad \text{by} \qquad (\alpha, \beta) \mapsto (\mathsf{head}(\alpha), (\mathsf{tail}(\alpha), \mathsf{tail}(\beta))).$$

Note that γ is well-defined since $(\mathsf{tail}(\alpha), \mathsf{tail}(\beta))$ is in R, because R is a bisimulation. Now it is straightforward to show that the two projection functions

$$\pi_1 \colon R \longrightarrow A^{\mathbb{N}} \qquad \text{and} \qquad \pi_2 \colon R \longrightarrow A^{\mathbb{N}}$$

are homomorphisms of coalgebras from (R, γ) to $(A^{\mathbb{N}}, \langle \mathsf{head}, \mathsf{tail} \rangle)$. Therefore it follows from the (uniqueness aspect of the) finality of $A^{\mathbb{N}}$ that $\pi_1 = \pi_2$. That is, if $R(\alpha, \beta)$ then $\alpha = \beta$.

We proceed by considering bisimulations for another functor.

Example 7.1 We define a functor B as

$$B(X) = \mathcal{P}(A \times X) = \{V \mid V \subseteq A \times X\}$$

As we saw in the fifth coalgebra example in Section 2, a labeled transition system $(X, A, \longrightarrow_X)$ can be identified with a B-coalgebra $\alpha_X : X \to B(X)$, where $(a, s') \in \alpha_X(s)$ if and only if $s \xrightarrow{a}_X s'$. In other words, the class of all labeled transition systems coincides with the class of all B-coalgebras. Let $(X, A, \longrightarrow_X)$ and $(Y, A, \longrightarrow_Y)$ be two labeled transition systems with the same set A of labels. An interesting question is what a coalgebra homomorphism between these two transition systems (as coalgebras α_X and α_Y) is, in terms of the transition structures $\longrightarrow_X$ and $\longrightarrow_Y$. Per definition, a B-homomorphism $f : \alpha_X \to \alpha_Y$ is a function $f : X \to Y$ such that $B(f) \circ \alpha_X = \alpha_Y \circ f$, where the function $B(f)$, also denoted by $\mathcal{P}(A \times f)$, is defined by

$$B(f)(V) = \mathcal{P}(A \times f)(V) = \{\langle a, f(s)\rangle \mid \langle a, s\rangle \in V\}.$$

One can easily prove that the equality $B(f) \circ \alpha_X = \alpha_Y \circ f$ is equivalent to the following two conditions:

1. for all s' in X, if $s \xrightarrow{a}_X s'$ then $f(s) \xrightarrow{a}_Y f(s')$;
2. for all t in Y, if $f(s) \xrightarrow{a}_Y t$ then there is s' in X with $s \xrightarrow{a}_X s'$ and $f(s') = t$.

Thus a homomorphism is a function that preserves and reflects transitions. This notion is quite standard, but sometimes only preservation is required, see e.g. [39].

There is the following well-known notion of bisimulation for transition systems [50, 53]: a bisimulation between between transition systems X and Y (as above) is a relation $R \subseteq X \times Y$ satisfying, for all $\langle s, t\rangle \in R$,

1. for all s' in X, if $s \xrightarrow{a}_X s'$ then there is t' in Y with $t \xrightarrow{a}_Y t'$ and $\langle s', t'\rangle \in R$;
2. for all t' in Y, if $t \xrightarrow{a}_Y t'$ then there is s' in X with $s \xrightarrow{a}_X s'$ and $\langle s', t'\rangle \in R$.

A concrete example of a bisimulation relation between two transition systems X and Y is the following. Consider two systems X and Y:

$$X = \left(\begin{array}{ccccc} s_0 & \xrightarrow{b} & s_1 & \xrightarrow{b} & \cdots \\ {\scriptstyle a}\downarrow & & {\scriptstyle a}\downarrow & & \\ s'_0 & & s'_1 & & \end{array} \right) \qquad Y = \left(\begin{array}{ccc} & & b \\ t & \circlearrowright & \\ {\scriptstyle a}\downarrow & & \\ t' & & \end{array} \right)$$

The relation $\{\langle s_i, s_j\rangle \mid i, j \geq 0\} \cup \{\langle s'_i, s'_j\rangle \mid i, j \geq 0\}$ is then a bisimulation on X. And $\{\langle s_i, t\rangle \mid i \geq 0\} \cup \{\langle s'_i, t'\rangle \mid i \geq 0\}$ is a bisimulation between X and Y. Note that the function $f : X \to Y$ defined by $f(s_i) = t$ and $f(s'_i) = t'$ is a homomorphism, and that there exists no homomorphism in the reverse direction from Y to X.

For cardinality reasons, a final B-coalgebra cannot exist: by Lemma 6.4 (ii), any final coalgebra is a fixed point: $X \cong \mathcal{P}(A \times X)$, and such a set does not exist because the cardinality of the latter set is strictly greater than that of X (for non-trivial sets of labels A). Therefore we restrict to so-called *finitely branching* transition systems, satisfying, for all states s,

$$\{\langle a, s\rangle \mid s \xrightarrow{a}_X s'\} \text{ is finite,}$$

Such systems can be identified with coalgebras of the functor

$$B_f(X) = \mathcal{P}_f(A \times X) = \{V \subseteq A \times X \mid V \text{ is finite}\}.$$

For this functor, a final coalgebra *does* exist. The proof, which is a bit technical, is due to Barr [8] (see also [64, 65]), and is omitted here (cf. the discussion on Section 2).

In what follows, let (P, π) be a final B_f-coalgebra. Borrowing the terminology of concurrency theory, we call the elements of P *processes*. As before, let $p \xrightarrow{a} p' \Leftrightarrow \langle a, p' \rangle \in \pi(p)$. We show that, similarly to the example of infinite lists, (P, π) satisfies a coinductive proof principle:

$$\text{if } R(p, p'), \text{ for some bisimulation } R \text{ on } P, \text{ then } p = p'. \tag{cpp}$$

The essence of the proof of this principle is again the observation that any bisimulation $R \subseteq P \times P$ can be supplied with a coalgebra structure: just define $\gamma \colon R \to \mathcal{P}_f(A \times R)$ on $\langle p, q \rangle$ in R, by

$$\gamma(\langle p, q \rangle) = \{\langle a, \langle p', q' \rangle\rangle \mid p \xrightarrow{a} p' \text{ and } q \xrightarrow{a} q'\}.$$

It follows from the bisimulation property of R that the two projection functions $\pi_1 \colon R \to P$ and $\pi_2 \colon R \to P$ are homomorphisms of coalgebras. Then by finality of P, both homomorphisms must be equal, which proves the cpp.

The present example is concluded by a definition, by coinduction, of a non-deterministic merge operation on processes, and a proof, by (cpp), of some of its properties. To this end, we supply $P \times P$ with the following B_f-coalgebra structure:

$$\begin{array}{ccc} (P \times P) & \xrightarrow{\quad\beta\quad} & \mathcal{P}_f(A \times (P \times P)) \\ \langle p, q \rangle & \longmapsto & \{\langle a, \langle p', q \rangle\rangle \mid p \xrightarrow{a} p'\} \cup \{\langle a, \langle p, q' \rangle\rangle \mid \ q \xrightarrow{a} q'\}. \end{array}$$

By finality of P, there exists a unique B_f-homomorphism $\mathsf{merge} \colon P \times P \to P$. It follows from the fact that merge is a homomorphism of transition systems (i.e. $\mathcal{P}(A \times \mathsf{merge}) \circ \beta = \pi \circ \mathsf{merge}$) that it satisfies the following rules:

$$\frac{p \xrightarrow{a} p'}{\mathsf{merge}\langle p, q \rangle \xrightarrow{a} \mathsf{merge}\langle p', q \rangle} \qquad \text{and} \qquad \frac{q \xrightarrow{a} q'}{\mathsf{merge}\langle p, q \rangle \xrightarrow{a} \mathsf{merge}\langle p, q' \rangle}.$$

The function merge satisfies a number of familiar properties. Let p_0 be the *terminated* process: formally, $p_0 = \pi^{-1}(\emptyset)$, for which no transitions exist. The following equalities

1. $\mathsf{merge}\langle p_0, p \rangle = p$;
2. $\mathsf{merge}\langle p, q \rangle = \mathsf{merge}\langle q, p \rangle$;
3. $\mathsf{merge}\langle \mathsf{merge}\langle p, q \rangle, r \rangle = \mathsf{merge}\langle p, \mathsf{merge}\langle q, r \rangle\rangle$,

are a consequence of (cpp) and the fact that the following relations are bisimulations on P:

1. $\{(\mathsf{merge}\langle p_0, p \rangle, p) \mid p \in P\}$;
2. $\{(\mathsf{merge}\langle p, q \rangle, \mathsf{merge}\langle q, p \rangle) \mid p, q \in P\}$;
3. $\{(\mathsf{merge}\langle \mathsf{merge}\langle p, q \rangle, r \rangle, \mathsf{merge}\langle p, \mathsf{merge}\langle q, r \rangle\rangle) \mid p, q, r \in P\}$.

For instance, the first relation is a bisimulation because we have transitions, for any p in P:

$$\mathsf{merge}\langle p_0, p \rangle \xrightarrow{a} \mathsf{merge}\langle p_0, p' \rangle \qquad \text{if and only if} \qquad p \xrightarrow{a} p',$$

and $(\mathsf{merge}\langle p_0, p' \rangle, p')$ is again in the relation. For the second relation, consider a pair of processes $(\mathsf{merge}\langle p, q \rangle, \mathsf{merge}\langle q, p \rangle)$, and suppose that we have a transition step

$$\mathsf{merge}\langle p, q \rangle \xrightarrow{a} r,$$

for some process r in P. (The other case, where a first step of $\mathsf{merge}\langle q, p \rangle$ is considered, is proved in exactly the same way.) It follows from the two rules above that one of the following two situations applies: either there exists a transition $p \xrightarrow{a} p'$ and $r = \mathsf{merge}\langle p', q \rangle$, or there exists a transition

$q \xrightarrow{a} q'$ and $r = \mathsf{merge}\langle p, q'\rangle$. Let us consider the first situation, the second being similar. If $p \xrightarrow{a} p'$ then it follows again from the rules above that there exists also a transition

$$\mathsf{merge}\langle q, p\rangle \xrightarrow{a} \mathsf{merge}\langle q, p'\rangle.$$

But then we have mimicked the transition step of $\mathsf{merge}\langle p, q\rangle$ by a transition step of $\mathsf{merge}\langle q, p\rangle$, in such a way that the resulting processes are again in the relation:

$$(\mathsf{merge}\langle p', q\rangle, \mathsf{merge}\langle q, p'\rangle)$$

is again a pair in relation 2. This shows that also relation 2 is a bisimulation. For 3, the same kind of argument can be given.

The notion of bisimulation was originally introduced in the semantics of concurrency [50, 53]. It has been used as a proof principle for final coalgebras by Aczel in his work on a theory of non-wellfounded sets [1]. Later a categorical definition of bisimulation was given in [2], by which bisimulation can be seen to be the coalgebraic dual of the notion of *congruence* on algebras, see the next section. The above definitions of a bisimulation (on infinite sequences $A^{\mathbb{N}}$ and finitely branching processes P) are special instances of this categorical definition, which is reproduced below.

Definition 7.2 Let T be a functor, and $c\colon U \to T(U)$ a T-coalgebra. A *bisimulation* on U is a relation R on U for which there exists a T-coalgebra structure $\gamma\colon R \to T(R)$ such that the two projection functions $\pi_1\colon R \to U$ and $\pi_2\colon R \to U$ are homomorphisms of T-coalgebras:

$$\begin{array}{ccccc} U & \xleftarrow{\pi_1} & R & \xrightarrow{\pi_2} & U \\ {\scriptstyle c}\downarrow & & \downarrow{\scriptstyle \gamma} & & \downarrow{\scriptstyle c} \\ T(U) & \xleftarrow[T(\pi_1)]{} & T(R) & \xrightarrow[T(\pi_2)]{} & T(U) \end{array}$$

The general formulation of the coinduction proof principle is then as follows.

Theorem 7.3 *Let $c\colon Z \xrightarrow{\cong} T(Z)$ be the final T-coalgebra. For all z and z' in P,*

$$\textit{if } R(z, z'), \textit{ for some bisimulation } R \textit{ on } Z, \textit{ then } z = z'. \qquad \text{(cpp)}$$

As in the examples above, the proof of this principle is immediate by finality. The reader is referred to [65] for further examples of definitions and proofs by coinduction. There exist other formalisations of the notion of bisimulation. In [28, 29] a bisimulation is described as a coalgebra in a category of relations, for a suitably lifted functor (associated with the original functor T). And in [39] bisimulations occur as suitable spans. But in a set-theoretic context, the above definition seems to be most convenient.

8 Predicates and relations on algebras and coalgebras

The logical arguments that we used so far involved:

- "Inductive" predicates on algebras for induction arguments. What we mean by "inductive" is that these predicates satisfy the induction assumptions for the underlying algebra. For example, an inductive predicate $P \subseteq \mathbb{N}$ on the natural numbers (see Example 5.4) satifies the assumptions $P(0)$ and $P(x) \Rightarrow P(S(x))$ for an induction argument. Such inductive predicates are suitably closed under the constructors of the algebra. One could also say: these predicates, as subsets, are subalgebras.

- Bisimulation relations on coalgebras. These are binary predicates which are closed under the destructors (or transitions) of the coalgebra. Or also, as subsets, they carry a coalgebra structure, as indicated in Definition 7.2. Such relations are used in coinduction arguments, of the following sort, see Theorem 7.3:

  Every bisimulation on a terminal coalgebra is contained in the equality relation.

Seeing this use of unary predicates on algebras and of binary predicates on coalgebras, one wonders if there is also a role for binary predicates on algebras and for unary predicates on coalgebras. The aim of this last section is to indicate that this is indeed the case.

A *congruence* relation R on an algebra is usually defined as a relation which is both (1) an equivalence relation, and (2) closed under the algebra's constructors. This second condition is equivalent to: R, as a subset, carries an algebra structure which makes both projection functions homomorphisms of algebras (dual to Definition 7.2, see also [64, 65]). Here we shall use "congruence" for a relation satisfying requirement (2), but not necessarily (1). Then we can formulate the *binary induction principle* as:

Every congruence relation on an initial algebra contains the equality relation.

Notice the perfect duality with the coinduction principle as formulated above. We note that a relation R contains the equality relation if and only if R is reflexive, i.e. if and only if $R(x, x)$ holds for all x.

For example, the binary induction principle on the natural numbers says that $R(x, x)$ holds for for a relation $R \subseteq \mathbb{N} \times \mathbb{N}$ if both $R(0, 0)$ and $R(x, y) \Rightarrow R(S(x), S(y))$ hold (for all $x, y \in \mathbb{N}$. It is easy to see that this binary induction principle is equivalent to the usual induction principle (on the natural numbers). This was first noted in [64], and proved subsequently for more general data types in [29] in an abstract categorical setting.

The situation for unary predicates on coalgebras is not fully settled yet. Of relevance are predicates which are closed under the coalgebra's destructors (or transitions). Equivalently, predicates which carry (as subsets) a subcoalgebra structure. They were first called "mongruences" in [32] (in analogy with congruences), but "invariants" [38] appears to be a better name, since if such a predicate holds for a state x, then it also holds for all successors of x. Such predicates are described as "subsystems" in [65].

Greatest invariants appear to be of interest: if P is a suitable property on a coalgebra, then the greatest invariant $\mathrm{inv}(P) \subseteq P$ is the final subcoalgebra in which P holds. This gives us a certain proof principle: if $Q \subseteq P$ and Q is an invariant, then $Q \subseteq \mathrm{inv}(P)$. Applications of this principle may be found in [32], where the final coalgebra satisfying certain constraints C is characterised as $\mathrm{inv}(C)$ (for C interpreted in the final coalgebra of the operations), and in [38] for proving the correctness of refinements between coalgebraic specifications (in an object-oriented setting).

The following table gives a summary of the predicates of interest on algebras and coalgebras.

	algebra	coalgebra
unary	inductive predicate (subalgebra)	invariant (subcoalgebra)
binary	congruence	bisimulation

Acknowledgements. Our thanks go to Ulrich Hensel, Marieke Huisman and Horst Reichel for their comments on an earlier version.

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