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Fibrational Induction Rules for Initial Algebras

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Abstract. This paper provides an induction rule that can be used to prove properties of data structures whose types are inductive, i.e., are carriers of initial algebras of functors. Our results are semantic in nature and are inspired by Hermida and Jacobs’ elegant algebraic formulation of induction for polynomial data types. Our contribution is to derive, under slightly different assumptions, an induction rule that is generic over all inductive types, polynomial or not. Our induction rule is generic over the kinds of properties to be proved as well: like Hermida and Jacobs, we work in a general fibrational setting and so can accommodate very general notions of properties on inductive types rather than just those of particular syntactic forms. We establish the correctness of our generic induction rule by reducing induction to iteration. We show how our rule can be instantiated to give induction rules for the data types of rose trees, finite hereditary sets, and hyperfunctions. The former lies outside the scope of Hermida and Jacobs’ work because it is not polynomial; as far as we are aware, no induction rules have been known to exist for the latter two in a general fibrational framework. Our instantiation for hyperfunctions underscores the value of working in the general fibrational setting since this data type cannot be interpreted as a set.

1 Introduction
Iteration operators provide a uniform way to express common and naturally occurring patterns of recursion over inductive data types. Expressing recursion via iteration operators makes code easier to read, write, and understand; facilitates code reuse; guarantees properties of programs such as totality and termination; and supports optimising program transformations such as fold fusion and short cut fusion. Categorically, iteration operators arise from initial algebra semantics of data types, in which each data type is regarded as the carrier of the initial algebra of a functor \( F \). Lambek’s Lemma ensures that this carrier is the least fixed point \( \mu F \) of \( F \), and initiality ensures that, given any \( F \)-algebra \( h : FA \to A \), there is a unique \( F \)-algebra homomorphism, denoted \( \text{fold}_h \), from the initial algebra in \( F(\mu F) \to \mu F \) to that algebra. For each functor \( F \), the map \( \text{fold} : (FA \to A) \to \mu F \to A \) is the iteration operator for the data type \( \mu F \). Initial algebra semantics thus provides a well-developed theory of iteration which is i) principled, and so helps ensure that programs have rigorous mathematical foundations that can be used to ascertain their meaning and correctness; ii) expressive, and so is applicable to all inductive types, rather than just syntactically defined classes of data types such as polynomial data types; and iii) correct, and so is valid in any model — set-theoretic, domain-theoretic, realizability, etc. — in which data types are interpreted as carriers of initial algebras.

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Because induction and iteration are closely linked we may reasonably expect that initial algebra semantics can be used to derive a principled, expressive, and correct theory of induction for data types we well. In most treatments of induction, given a functor \( F \) together with a property \( P \) to be proved about data of type \( \mu F \), the premises of the induction rule for \( \mu F \) constitute an \( F \)-algebra with carrier \( \Sigma x : \mu F. P x \). The conclusion of the rule is obtained by supplying such an \( F \)-algebra as input to the \textit{fold} for \( \mu F \). This yields a function from \( \mu F \) to \( \Sigma x : \mu F. P x \) from which function of type \( \forall x : \mu F. P x \) can be obtained. It has not, however, been possible to characterise \( F \)-algebras with carrier \( \Sigma x : \mu F. P x \) without additional assumptions on \( F \). Induction rules are thus typically derived under the assumption that the functors involved have a certain structure, e.g., that they are polynomial. Moreover, taking the carriers of the algebras to be \( \Sigma \)-types assumes that properties are represented as type-valued functions. So while induction rules derived as described are both principled and correct, their expressiveness is limited along two dimensions: with respect to the data types for which they can be derived and the nature of the properties they can verify.

One principled and correct approach to induction is given by Hermida and Jacobs [6]. They lift each functor \( F \) on a base category of types to a functor \( \hat{F} \) on a category of properties over those types, and take the premises of the induction rule for the type \( \mu F \) to be an \( \hat{F} \)-algebra. Hermida and Jacobs work in a fibrational setting and the notion of property they consider is, accordingly, very general. Indeed, they accommodate any notion of property that can be fibred over the category of types, and so overcome one of the two limitations mentioned above. On the other hand, their approach is only applicable to polynomial data types, so the limitation on the class of data types treated remains in their work.

This paper shows how to remove the restriction on the class of data types treated. \textit{Our main result is a derivation of a generic induction rule that can be instantiated to every inductive type} — i.e., to every type which is the carrier of the initial algebra of a functor — regardless of whether it is polynomial. We take Hermida and Jacobs’ approach as our point of departure and show that, under slightly different assumptions on the fibration involved, we can lift \textit{any} functor on its base category that has an initial algebra to a functor on its properties. This is clearly an important theoretical result, but it also has practical consequences:

- We show in Example 2 how our generic induction rule can be instantiated to the codomain fibration to derive the rule for rose trees that one would intuitively expect. The data type of rose trees lies outside the scope of Hermida and Jacobs’ results because it is not polynomial. On the other hand, an induction rule for rose trees is available in the proof assistant Coq, although it is neither the one we intuitively expect nor expressive enough to prove properties that ought to be amenable to inductive proof. The rule we derive for rose trees is indeed the expected one, which suggests that our derivation may enable automatic generation of more useful induction rules in Coq, rather than requiring the user to hand code them as is currently necessary.
- We further show in Example 3 how our generic induction rule can be instantiated, again to the codomain fibration, to derive a rule for the data type of
finite hereditary sets. This data type is defined in terms of quotients and so lies outside current theories of induction.

Finally, we show in Example 4 how our generic induction rule can be instantiated to the subobject fibration over \( \omega \text{cpo} \) to derive a rule for the data type of hyperfunctions. Because this data type cannot be interpreted as a set, a fibration other than the codomain fibration over \( \text{Set} \) is required; in this case, use of the subobject fibration allows us to derive an induction rule for admissible subsets of hyperfunctions. Moreover, the functor underlying the data type of hyperfunctions is not strictly positive, and this fact again underscores the advantage of being able to handle a very general class of functors. As far as we know, induction rules for finite hereditary sets and hyperfunctions have not previously existed in the general fibrational framework.

Although our theory of induction is applicable to all functors having initial algebras, including higher-order ones, our examples show that working in the general fibrational setting is beneficial even if attention is restricted to first-order functors. Note also that our induction rules coincide with those of Hermida and Jacobs when specialised to polynomial functors in the codomain fibration. But the structure we require of fibrations generally is slightly different from that required by Hermida and Jacobs, so while our theory is in essence a generalisation of theirs, the two are, strictly speaking, incomparable. The structure we require is, however, still minimal and certainly present in all standard fibrational models of type theory (see Section 4). Like Hermida and Jacobs, we prove our generic induction rule correct by reducing induction to iteration.

We take a purely categorical approach to induction in this paper, and derive our generic induction rule from only the initial algebra semantics of data types. As a result, our work is inherently extensional. While translating our constructions into intensional settings may therefore require additional effort, we expect the guidance offered by the categorical viewpoint to support the derivation of induction rules for functors that are not treatable at present. Since we do not use any form of impredicativity in our constructions, and instead use only the weaker assumption that initial algebras exist, this guidance will be widely applicable.

The remainder of this paper is structured as follows. To make our results as accessible as possible, we illustrate them in Section 2 with a categorical derivation of the familiar induction rule for the natural numbers. In Section 3 we derive an induction rule for the special case of the codomain fibration over \( \text{Set} \), i.e., for functors on the category of sets and properties representable as set-valued predicates. We also show how this rule can be instantiated to derive the one from Section 2, and the ones for rose trees and finite hereditary sets mentioned above. Then, in Section 4 we sketch the general fibrational form of our derivation (space constraints prevent a full treatment) and illustrate the resulting generic induction rule with the aforementioned application to hyperfunctions. Section 5 concludes and offers some additional directions for future research.

When convenient, we identify isomorphic objects of a category and write \( = \) rather than \( \simeq \). We write \( 1 \) for the canonical singleton set and denote its single element by \( \cdot \). In Sections 2 and 3 we assume that types are interpreted as objects in the category \( \text{Set} \) of sets, and so \( 1 \) also denotes the unit type in those sections.
2 A Familiar Induction Rule

Consider the inductive data type Nat, which defines the natural numbers and can be specified in a programming language with Haskell-like syntax by

\[
data Nat = \text{Zero} | \text{Succ Nat}
\]

The observation that Nat is the least fixed point of the functor \(N\) on Set defined by \(NX = 1 + X\) can be used to define the following iteration operator for it:

\[
foldNat : X \rightarrow (X \rightarrow X) \rightarrow \text{Nat} \rightarrow X
\]

\[
foldNat z s \text{Zero} = z
\]

\[
foldNat z s \text{(Succ n)} = s (foldNat z s n)
\]

Categorically, iteration operators such as foldNat arise from the initial algebra semantics of data types. If \(B\) is a category and \(F\) is a functor on \(B\), then an \(F\)-algebra is a morphism \(h : FX \rightarrow X\) for some object \(X\) of \(B\). We call \(X\) the carrier of \(h\). For any functor \(F\), the collection of \(F\)-algebras itself forms a category \(\text{Alg}_F\) which we call the category of \(F\)-algebras. In \(\text{Alg}_F\), an \(F\)-algebra morphism between \(F\)-algebras \(h : FX \rightarrow X\) and \(g : FY \rightarrow Y\) is a map \(f : X \rightarrow Y\) such that \(f \circ h = g \circ Ff\). When it exists, the initial \(F\)-algebra \(\text{in} : F(\mu F) \rightarrow \mu F\) is unique up to isomorphism and has the least fixed point \(\mu F\) of \(F\) as its carrier. Initiality ensures that there is a unique \(F\)-algebra morphism \(\text{fold} h : \mu F \rightarrow X\) from \(\text{in}\) to any other \(F\)-algebra \(h : FX \rightarrow X\). This gives rise to the following iteration operator \(\text{fold}\) for \(F\) or, equivalently, for the inductive type \(\mu F\):

\[
\text{fold} : (FX \rightarrow X) \rightarrow \mu F \rightarrow X
\]

\[
\text{fold} h (\text{in} t) = h (F (\text{fold} h) t)
\]

Since \(\text{fold}\) is derived from initial algebra semantics it is principled and correct. It is also expressive, since it can be defined for every inductive type. In fact, \(\text{fold}\) is a single iteration operator parameterised over inductive types rather than a family of iteration operators, one for each such type, and the iteration operator \(\text{foldNat}\) above is the instantiation to \(\text{Nat}\) of the generic iteration operator \(\text{fold}\).

The iteration operator \(\text{foldNat}\) can be used to derive the standard induction rule for \(\text{Nat}\). This rule says that if a property \(P\) holds for \(0\), and if \(P\) holds for \(n+1\) whenever it holds for a natural number \(n\), then \(P\) holds for all natural numbers. Representing each property of natural numbers as a predicate \(P : \text{Nat} \rightarrow \text{Set}\) mapping each \(n : \text{Nat}\) to the set of proofs that \(P\) holds for \(n\), we wish to represent this rule at the object level as a function \(\text{indNat}\) with type

\[
\forall(P : \text{Nat} \rightarrow \text{Set}). P \text{Zero} \rightarrow (\forall n : \text{Nat}. P n \rightarrow P (\text{Succ} n)) \rightarrow (\forall n : \text{Nat}. P n)
\]

Code fragments such as the above, which involve quantification over sets, properties, or functors, are to be treated as “categorically inspired” within this paper. This is because quantification over such higher-kindred objects cannot be interpreted in \(\text{Set}\). In order to give a formal interpretation to code fragments like the one above, we would need to work in a category such as that of modest sets. The ability to work with functors over categories other than \(\text{Set}\) is one of the
motivations for working in the general fibrational setting of Section 4. Of course, the use of category theory to suggest computational constructions has long been accepted within the functional programming community (see, e.g., [1, 2, 13]).

A function \( \text{indNat} \) with the above type takes as input the property \( P \) to be proved, a proof \( \phi \) that \( P \) holds for \( \text{Zero} \), and a function \( \psi \) mapping each \( n : \text{Nat} \) and each proof that \( P \) holds for \( n \) to a proof that \( P \) holds for \( \text{Succ} \ n \), and returns a function mapping each \( n : \text{Nat} \) to a proof that \( P \) holds for \( n \). We can write \( \text{indNat} \) in terms of \( \text{foldNat} \) — and thus reduce induction for \( \text{Nat} \) to iteration for \( \text{Nat} \) — as follows. First note that \( \text{indNat} \) cannot be obtained by instantiating the type \( X \) in the type of \( \text{foldNat} \) to a type of the form \( P_n \) for a specific \( n \) because \( \text{indNat} \) returns elements of the types \( P_n \) for different values \( n \) and these types are, in general, distinct from one another. We therefore need a type containing all of the elements of \( P_n \) for every \( n \). Such a type can be formally given by the dependent type \( \Sigma n : \text{Nat}. P_n \) comprising pairs \((n, p)\) where \( n : \text{Nat} \) and \( p : P_n \).

The standard approach to defining \( \text{indNat} \) is thus to apply \( \text{foldNat} \) to an \( \text{N}-\)algebra with carrier \( \Sigma n : \text{Nat}. P_n \). Such an algebra has components \( \alpha : \Sigma n : \text{Nat}. P_n \) and \( \beta : \Sigma n : \text{Nat}. P_n \to \Sigma n : \text{Nat}. P_n \). Given \( \phi : P \text{Zero} \) and \( \psi : \forall n. P_n \to P(\text{Succ} \ n) \), we choose \( \alpha = (\text{Zero}, \phi) \) and \( \beta(n, p) = (\text{Succ} \ n, \psi n p) \) and note that \( \text{foldNat} \alpha \beta : \text{Nat} \to \Sigma n : \text{Nat}. P_n \). We tentatively take \( \text{indNat} P \phi \psi n \) to be \( p \), where \( \text{foldNat} \alpha \beta n = (m, p) \). But in order to know that \( p \) actually gives a proof for \( n \) itself, we must show that \( m = n \). Fortunately, this follows easily from the uniqueness of \( \text{foldNat} \alpha \beta \). Letting \( \pi_P \) be the second projection on dependent pairs, the induction rule for \( \text{Nat} \) is

\[
\text{indNat} : \forall(P : \text{Nat} \to \text{Set}). P \text{Zero} \to (\forall n : \text{Nat}. P_n \to P(\text{Succ} \ n)) \\
\to (\forall n : \text{Nat}. P_n) \\
\text{indNat} P \phi \psi = \pi_P \circ (\text{foldNat}(\text{Zero}, \phi) (\lambda(n, p). \text{Succ} \ n, \psi n p))
\]

The use of dependent types is fundamental to this formalization of the induction rule for \( \text{Nat} \), but this is only possible because properties are taken to be set-valued functions. The remainder of this paper uses fibrations to generalise the above treatment of induction to arbitrary functors which have initial algebras and arbitrary properties which are fibred over the category whose objects interpret types. In the general fibrational setting, properties are given axiomatically via the fibrational structure rather than assumed to be (set-valued) functions.

3 Induction rules over Set

The main result of this paper is the derivation of an induction rule that is generic over inductive types and can be used to verify any notion of property that is fibred over the category whose objects interpret types. In the remainder of the paper we restrict attention to functors that have initial algebras. In this section we further assume that types are modelled by sets, so the functors we consider are on \( \text{Set} \) and the properties we consider are functions mapping data to sets of proofs that these properties hold for them. We make these assumptions to present our derivation in the simplest setting possible. But they are not always valid, so we derive a more general induction rule which relaxes them in Section 4. It should be more easily digestible once the derivation in this section is understood.
The derivation for $Nat$ in Section 2 suggests that an induction rule for an inductive data type $\mu F$ should, in general, look something like this:

$$\text{ind} : \forall P : \mu F \rightarrow \text{Set} . ??? \rightarrow \forall x : \mu F. P x$$

But what should the premises — denoted ??? here — of the generic induction rule ind be? Since we want to construct, for any term $x : \mu F$, a proof term of type $Px$ from proof terms for $x$’s substructures, and since the functionality of the fold operator for $\mu F$ is precisely to compute a value for $x : \mu F$ from the values for $x$’s substructures, it is natural to try to equip $P$ with an $F$-algebra structure that can be input to fold to yield a mapping of each $x : \mu F$ to an element of $Px$. But this approach quickly hits a snag. Since the codomain of every predicate $P : \mu F \rightarrow \text{Set}$ is Set itself, rather than an object of Set, $F$ cannot be applied to $P$ as is needed to equip $P$ with an $F$-algebra structure. Moreover, an induction rule for $\mu F$ cannot be obtained by applying fold to an $F$-algebra with carrier $Px$ for any specific $x$.

Such considerations led Hermida and Jacobs [6] to define a category of predicates $\mathcal{P}$ and a lifting $\hat{F}$ for every polynomial functor $F$ on Set to a functor $\hat{F}$ on $\mathcal{P}$ that respects the structure of $F$. They then constructed $\hat{F}$-algebras with carrier $P x$ for any specific $x$.

Definition 1 Let $X$ be a set. A predicate on $X$ is a pair $(X, P)$ where $P : X \rightarrow \text{Set}$ maps each $x \in X$ to a set $Px$. We call $X$ the domain of the predicate $(X, P)$.

Definition 2 The category of predicates $\mathcal{P}$ has predicates as its objects. A morphism from a predicate $(X, P)$ to a predicate $(X', P')$ is a pair $(f, f^-) : (X, P) \rightarrow (X', P')$ of functions, where $f : X \rightarrow X'$ and $f^- : \forall x : X. P x \rightarrow P'(fx)$.

The notion of a morphism from $(X, P)$ to $(X', P')$ does not require the sets of proofs $Px$ and $P'(fx)$, for any $x \in X$, to be equal. Instead, it requires only the existence of a function $f^-$ which maps, for each $x$, each proof in $Px$ to a proof in $P'(fx)$. We denote by $U : \mathcal{P} \rightarrow \text{Set}$ the forgetful functor mapping each predicate $(X, P)$ to its domain $X$ and each predicate morphism $(f, f^-)$ to $f$.

An alternative to Definition 2 would take the category of predicates to be the arrow category over Set, but the natural lifting in this setting does not generalise to arbitrary fibrations. Indeed, if properties are modelled as functions, then every functor can be applied to a property, and hence every functor can be its own lifting. In the general fibrational setting, properties are not necessarily modelled by functions, so a functor cannot, in general, be its own lifting. The decision not
to use arrow categories to model properties is thus dictated by our desire to lift functors in such a way that it can be extended to the general fibrational setting.

**Definition 3** Let $F$ be a functor on $\mathcal{Set}$. A lifting of $F$ from $\mathcal{Set}$ to $\mathcal{P}$ is a functor $\hat{F}$ on $\mathcal{P}$ such that $FU = U\hat{F}$.

Note that if $P$ is a predicate on $X$, then $\hat{F}P$ is a predicate on $FX$. We can now derive the standard induction rule from Section 2 for $\mathcal{Nat}$ as follows.

**Example 1** The data type of natural numbers is $\mu N$ where $N$ is the functor on $\mathcal{Set}$ defined by $NX = 1 + X$. If $P$ is a predicate on $X$, then a lifting $NP : 1 + X \rightarrow X$ of $N$ from $\mathcal{Set}$ to $\mathcal{P}$ is given by $NP (\text{id}) = 1$ and $NP (\text{inr} n) = Pn$. An $\hat{N}$-algebra with carrier $P : \mathcal{Nat} \rightarrow \mathcal{Set}$ can be given as follows. For every predicate $P$ over $\mathcal{Nat}$ and a function $h : \forall n : \mathcal{Nat}. Pn \rightarrow P(n+1)$, the second component defined by ${\forall n : \mathcal{Nat.} Pn \rightarrow P(n+1)}$ in the lifting $\hat{N}$ of the $\hat{N}$-algebra with carrier $P : \mathcal{Nat} \rightarrow \mathcal{Set}$ and first component in gives the premises of the familiar induction rule in Example 1.

The notion of predicate comprehension is a key ingredient of our lifting.

**Definition 4** Let $P$ be a predicate on $X$. The comprehension of $P$, denoted $\{ P \}$, is the type $\Sigma x : X. Px$ comprising pairs $(x, p)$ where $x : X$ and $p : Px$. The map taking each predicate $P$ to $\{ P \}$, and taking each predicate morphism $(f, f^\sim) : P \rightarrow P'$ to $\{ (f, f^\sim) \} : \{ P \} \rightarrow \{ P' \}$ defined by $\{ (f, f^\sim) \}(x, p) = (fx, f^\sim x p)$, defines the comprehension functor $\{ - \}$ from $\mathcal{P}$ to $\mathcal{Set}$.

**Definition 5** If $F$ is a functor on $\mathcal{Set}$, then the lifting $\hat{F}$ is the functor on $\mathcal{P}$ given as follows. For every predicate $P$ on $X$, $FP : FX \rightarrow \mathcal{Set}$ is defined by $FP = (F \pi_P)^{-1}$, where the natural transformation $\pi : \{ - \} \rightarrow U$ is given by $\pi_P(x, p) = x$. For every predicate morphism $f : P \rightarrow P'$, $\hat{F}f = (k, k^\sim)$ where $k = F(\text{id})$, and $k^\sim : \forall y : FX. \hat{F}Py \rightarrow \hat{F}P'(ky)$ is given by $k^\sim y z = F\{ f \}z$.

The inverse image $f^{-1} f : X \rightarrow Y$ is a predicate $P : Y \rightarrow \mathcal{Set}$. Thus if $P$ is a predicate on $X$, then $\hat{F}P$ is a predicate on $FX$, so $\hat{F}$ is a lifting of $F$ from $\mathcal{Set}$ to $\mathcal{P}$. The lifting $\hat{F}$ captures an “all” modality generalising Haskell’s `all` function on lists to arbitrary data types. A similar modality is given in [12] for indexed containers.

The lifting in Example 1 is the instantiation of the construction in Definition 5 to $NX = 1 + X$ on $\mathcal{Set}$. Indeed, if $P$ is any predicate, then $\hat{N}P = (N \pi_P)^{-1}$, i.e., $\hat{N}P = (\text{id} + \pi_P)^{-1}$, by Definition 5. Since the inverse image of the coproduct of functions is the coproduct of their inverse images, since $id^{-1} 1 = 1$, and since $\pi_P^{-1} n = \{(n, p) : Pn\}$ for all $n$, we have $\hat{N}P (\text{id}) = 1$ and $\hat{N}P (\text{inr} n) = Pn$.

The rest of this section shows that $\hat{F}$-algebras with carrier $\{ P \}$ are inter-derivable with $\hat{F}$-algebras with carrier $P$, and then uses this result to derive our induction rule.

**Definition 6** The functor $K_1 : \mathcal{Set} \rightarrow \mathcal{P}$ maps each set $X$ to the predicate $K_1X = \lambda x : X. 1$ and each $f : X \rightarrow Y$ to the predicate morphism $(f, \lambda x : X. id)$. 
The predicate $K_1 X$ is called the truth predicate on $X$. For every $x : X$, the set $K_1 X$ of proofs that $K_1 X$ holds for $x$ is a singleton, and thus is non-empty. For any functor $F$, the lifting $\hat{F}$ maps the truth predicate on a set $X$ to that on $FX$.

**Lemma 1** For any functor $F$ on $\text{Set}$ and any $X : \text{Set}$, $\hat{F}(K_1 X) = K_1(FX)$.

**Proof:** By Definition 5, $\hat{F}(K_1 X) = (F\pi_{K_1 X})^{-1}$. We have that $\pi_{K_1 X}$ is a isomorphism since there is only one proof of $K_1 X$ for each $x : X$, and thus that $F\pi_{K_1 X}$ is an isomorphism as well. As a result, $(F\pi_{K_1 X})^{-1}$ maps every $y : FX$ to a singleton set, and therefore $\hat{F}(K_1 X) = (F\pi_{K_1 X})^{-1} = \lambda y : FX. 1 = K_1(FX)$.

The fact that $K_1 \dashv \{\{\}\}$ is critical to the constructions below. This is proved in [6]; we include its proof here for completeness and to establish notation. The description of comprehension as a right adjoint can be traced back to Lawvere [10].

**Lemma 2** $K_1$ is left adjoint to $\{\{\}\}$.

**Proof:** We must show that, for any predicate $P$ and any set $Y$, the set $\mathcal{P}(K_1 Y, P)$ of morphisms from $K_1 Y$ to $P$ in $\mathcal{P}$ is in bijective correspondence with the set $\text{Set}(Y, \{P\})$ of morphisms from $Y$ to $\{P\}$. Define maps $(-)^\dagger : \text{Set}(Y, \{P\}) \rightarrow \mathcal{P}(K_1 Y, P)$ and $(-)^\# : \mathcal{P}(K_1 Y, P) \rightarrow \text{Set}(Y, \{P\})$ by $h^\dagger = (h_1, h_2)$ where $h_y = (v, p)$, $h_1 y = v$ and $h_2 y = p$, and $(k, k^{-}\#) = \lambda (y : Y). (ky, k^{-} y)$. These give a natural isomorphism between $\text{Set}(Y, \{P\})$ and $\mathcal{P}(K_1 Y, P)$.

Naturality of $(-)^\dagger$ ensures that $(g \circ f)^\dagger = g^\dagger \circ K_1 f$ for all $f : Y' \rightarrow Y$ and $g : Y \rightarrow \{P\}$. Similarly for $(-)^\#$. Moreover, $id^\dagger$ is the counit of the adjunction between $K_1$ and $\{\{\}\}$. These observations are used in the proof of Lemma 4.

**Lemma 3** There is a functor $\Phi : \text{Alg}_F \rightarrow \text{Alg}_{\hat{F}}$ such that if $k : FX \rightarrow X$, then $\Phi k : \hat{F}(K_1 X) \rightarrow K_1 X$.

**Proof:** For an $F$-algebra $k : FX \rightarrow X$ define $\Phi k = K_1 k$, and for two $F$-algebras $k : FX \rightarrow X$ and $k' : FX' \rightarrow X'$ and an $F$-algebra morphism $h : X \rightarrow X'$ between them define the $\hat{F}$-algebra morphism $\Phi h : \Phi k \rightarrow \Phi k'$ by $\Phi h = K_1 h$. Then $K_1(FX) = \hat{F}(K_1 X)$ by Lemma 1, so that $\Phi k$ is an $\hat{F}$-algebra and $K_1 h$ is an $F$-algebra morphism. It is easy to see that $\Phi$ preserves identities and composition.

**Lemma 4** The functor $\Phi$ has a right adjoint $\Psi$ such that if $j : \hat{F} P \rightarrow P$, then $\Psi j : F\{P\} \rightarrow \{P\}$.

**Proof:** We construct $\Psi : \text{Alg}_{\hat{F}} \rightarrow \text{Alg}_F$ as follows. Given an $\hat{F}$-algebra $j : \hat{F} P \rightarrow P$, we use the fact that $\hat{F}(K_1 \{P\}) = K_1(F\{P\})$ by Lemma 1 to define $\Psi j : F\{P\} \rightarrow \{P\}$ by $\Psi j = (j \circ \hat{F} id)^\dagger$. To specify the action of $\Psi$ on an $\hat{F}$-algebra morphism $h$, define $\Psi h = \{h\}$. Clearly $\Psi$ preserves identity and composition.

Next we show $\Phi \dashv \Psi$, i.e., for every $F$-algebra $k : FX \rightarrow X$ and $\hat{F}$-algebra $j : \hat{F} P \rightarrow P$ with $P$ a predicate on $X$, there is a natural isomorphism between $F$-algebra morphisms from $k$ to $\Psi j$ and $\hat{F}$-algebra morphisms from $\Phi k$ to $j$. First
observe that an $F$-algebra morphism from $k$ to $\Psi j$ is a map from $X$ to $\{P\}$, and an $\hat{F}$-algebra morphism from $\Phi k$ to $j$ is a map from $K_1 X \to P$. An isomorphism between such maps is given by the adjunction $K_1 \dashv \{-\}$ from Lemma 2, and so is natural. We must check that $f : X \to \{P\}$ is an $F$-algebra morphism from $k$ to $\Psi j$ iff $f^\dagger : K_1 X \to P$ is an $\hat{F}$-algebra morphism from $\Phi k$ to $j$.

So assume $f : X \to \{P\}$ is an $F$-algebra morphism from $k$ to $\Psi j$, i.e., $f \circ k = \Psi j \circ F f$. We must prove that $f^\dagger \circ \Phi k = j \circ \hat{F} f^\dagger$. By the definition of $\Phi$ in Lemma 3, thus amounts to showing $f^\dagger \circ K_1 k = j \circ \hat{F} f^\dagger$. Since $(-)^\dagger$ is an isomorphism, $f$ is an $F$-algebra morphism iff $(f \circ k)^\dagger = (\Psi j \circ F f)^\dagger$. Naturality of $(-)^\dagger$ ensures that $(f \circ k)^\dagger = f^\dagger \circ K_1 k$ and that $(\Psi j \circ F f)^\dagger = (\Psi j)^\dagger \circ K_1 (F f)$, so the previous equality holds iff $f^\dagger \circ K_1 k = (\Psi j)^\dagger \circ K_1 (F f)$. But

\[
\begin{align*}
j \circ \hat{F} f^\dagger &= j \circ \hat{F} (id^\dagger \circ K_1 f) & \text{by naturality of } (-)^\dagger \text{ and } f = id \circ f \\
&= (j \circ \hat{F} id^\dagger) \circ \hat{F} (K_1 f) & \text{by the functoriality of } \hat{F} \\
&= (\Psi j)^\dagger \circ K_1 (F f) & \text{by the definition of } \Psi, \text{ the fact that } (-)^\dagger \text{ and } (-)^\# \\
&= f^\star \circ K_1 k & \text{are inverses, and Lemma 1} \\
&= f^\dagger \circ \Phi k & \text{by the observation immediately preceding this proof} \\
&= f^\dagger \circ K_1 k & \text{by the definition of } \Phi
\end{align*}
\]

So $f^\dagger \circ K_1 k = (\Psi j)^\dagger \circ K_1 (F f)$, and $f^\dagger$ is an $\hat{F}$-algebra morphism from $\Phi k$ to $j$.

Lemma 4 ensures that $F$-algebras with carrier $\{P\}$ are interderivable with $\hat{F}$-algebras with carrier $P$. For example, the $N$-algebra $[\alpha, \beta]$ with carrier $\{P\}$ from Section 2 can be derived from the $\hat{N}$-algebra with carrier $P$ given in Example 1. Since we define a lifting $\hat{F}$ for any functor $F$, Lemma 4 thus shows how to construct $F$-algebras with carrier $\Sigma x : \mu F. P x$ for any $F$.

We can now derive our generic induction rule. For every predicate $P$ on $X$ and every $F$-algebra $(k, k^-) : \hat{F} P \to P$, Lemma 4 ensures that $\Psi$ constructs from $(k, k^-)$ an $F$-algebra with carrier $\{P\}$. Thus, $\text{fold } (\Psi (k, k^-)) : \mu F \to \{P\}$ and this map decomposes into two parts: $\phi = \pi_P \circ \text{fold } (\Psi (k, k^-)) : \mu F \to X$ and $\psi : \forall (t : \mu F). P (\phi t)$. Initiality of $\text{in}$ ensures $\phi = \text{fold } k$. This gives the following generic induction rule for the type $X$, which reduces induction to iteration:

\[
\begin{align*}
genind & : \forall (F : \text{Set} \to \text{Set}) (P : X \to \text{Set}) ((k, k^-) : (\hat{F} P \to P)) (x : \mu F). P (\text{fold } k x) \\
genind P F & = \pi_P \circ \text{fold } \Psi
\end{align*}
\]

When $X = \mu F$ and $k = \text{in}$, initiality of $\text{in}$ further ensures that $\phi = \text{fold } \text{in} = \text{id}$, and thus that $\text{genind}$ specialises to the expected induction rule for an inductive data type $\mu F$:

\[
\begin{align*}
\text{ind} & : \forall (F : \text{Set} \to \text{Set}) (P : \mu F \to \text{Set}) ((k, k^-) : (\hat{F} P \to P)). (k = \text{in}) \to \forall (x : \mu F). P x \\
\text{ind } F P & = \pi_P \circ \text{fold } \Psi
\end{align*}
\]
This rule can be instantiated to familiar rules for polynomial data types, as well as to ones we would expect for data types such as rose trees and finite hereditary sets, both of which lie outside the scope of standard methods.

**Example 2** The data type of rose trees is given in Haskell-like syntax by

```
data Rose = Node(Int, List Rose)
```

The functor underlying Rose is $F X \equiv \text{Int} \times \text{List } X$ and its induction rule is

\[
\text{indRose} \quad : \forall (P : \text{Rose} \to \text{Set}) \to (k, k^\sim) \to (\hat{F} P \to P).
\]

\[
(k = \text{in}) \to \forall (x : \text{Rose}). P x
\]

\[
\text{indRose } \hat{F} P = \pi_\p \circ \text{fold } \circ \Psi
\]

Calculating $\hat{F} P = (F \pi P)^{-1} : F \text{Rose } \to \text{Set}$, and writing $xs !! k$ for the $k^{th}$ component of a list $xs$, we have that

\[
\hat{F} P (i, rs)
\]

\[
= \{ z : F\{P\} \mid F_\pi P z = (i, rs) \}
\]

\[
= \{ (j, cps) : \text{Int } \times \text{List } \{P\} \mid F_\pi P (j, cps) = (i, rs) \}
\]

\[
= \{ (j, cps) : \text{Int } \times \text{List } \{P\} \mid (id \times \text{List } \pi P)(j, cps) = (i, rs) \}
\]

\[
= \{ (j, cps) : \text{Int } \times \text{List } \{P\} \mid j = i \text{ and } \pi P \text{ cps } = rs \}
\]

\[
= \{ (j, cps) : \text{Int } \times \text{List } \{P\} \mid j = i \text{ and } \forall k < \text{length } \pi P \text{ (cps }!! k) = rs !! k \}
\]

An $\hat{F}$-algebra whose underlying $F$-algebra is in : $F \text{Rose } \to \text{Rose}$ is thus a pair of functions $(i, k^\sim)$, where $k^\sim$ has type

\[
\forall i : \text{Int}. \forall rs : \text{List Rose},
\]

\[
\{ (j, cps) : \text{Int } \times \text{List } \{P\} \mid j = i \text{ and } \forall k < \text{length } \pi P \text{ (cps }!! k) = rs !! k \}
\]

\[
\to P(\text{Node } (i, rs))
\]

\[
= \forall i : \text{Int}. \forall rs : \text{List Rose},
\]

\[
\{ \text{cps} : \text{List } \{P\} \mid \forall k < \text{length } \pi P \text{ (cps }!! k) = rs !! k \}
\]

\[
\to P(\text{Node } (i, rs))
\]

\[
= \forall i : \text{Int}. \forall rs : \text{List Rose}. (\forall k < \text{length } rs. P (rs !! k)) \to P(\text{Node } (i, rs))
\]

The last equality is due to surjective pairing for dependent products and the fact that $\text{length } \pi P = \text{length } rs$. The type of $k^\sim$ gives the hypotheses of the induction rule for rose trees.

**Example 3** Hereditary sets are sets whose elements are themselves sets, and so are the core data structures within set theory. The data type $\text{HS}$ of finite hereditary sets is $\mu P_f$ for the finite powerset functor $P_f$. If $P : X \to \text{Set}$, then $P_f \pi P : P_f(\Sigma x : X. P x) \to P_f X$ maps each set $\{(x_1, p_1), \ldots, (x_n, p_n)\}$ to the set $\{x_1, \ldots, x_n\}$, so that $(P_f \pi P)^{-1}$ maps a set $\{x_1, \ldots, x_n\}$ to the set $P x_1 \times \ldots \times P x_n$. A $P_f$-algebra with carrier $P : \text{HS } \to \text{Set}$ and first component in therefore has as its second component a function of type

\[
\forall \{s_1, \ldots, s_n\} : P_f(\text{HS}). P s_1 \times \ldots \times P s_n \to P(\text{in } \{s_1, \ldots, s_n\})
\]

The induction rule for finite hereditary sets is thus

\[
\text{indHS} \quad : \forall \{s_1, \ldots, s_n\} : P_f(\text{HS}). P s_1 \times \ldots \times P s_n \to P(\text{in } \{s_1, \ldots, s_n\})
\]

\[
\to \forall (s : \text{HS}). P s
\]
4 Induction rules in the fibrational setting

We can treat a more general notion of predicate using fibrations. We motivate the move from the codomain fibration to arbitrary fibrations by observing that i) the semantics of data types in languages involving recursion and other effects usually involves categories other than \( \text{Set} \); ii) in such circumstances, the notion of a predicate can no longer be taken as a function with codomain \( \text{Set} \); iii) even when working in \( \text{Set} \) there are reasonable notions of “predicate” other than that in Section 3 (for example, a predicate on a set \( X \) could be a subobject of \( X \)); and iv) when, in future work, we come to consider induction rules for data types such as nested types, GADTs, indexed containers, and dependent types (see Section 5), we will want to appropriately instantiate a general theory of induction rather than having to invent a new one. Thus, although we could develop an ad hoc theory for each choice of category, functor, and predicate, it is far preferable to develop a uniform, axiomatic approach that is widely applicable.

Fibrations support precisely this kind of axiomatic approach, so this section generalises the constructions of the previous one to the general fibrational setting. The standard model of type theory based on locally cartesian closed categories does arise as a specific fibration — namely, the codomain fibration — but the general fibrational setting is far more flexible. In locally cartesian closed models of type theory, predicates and types coexist in the same category, so a functor can be taken to be its own lifting. In the general setting, predicates are not simply functions or morphisms, properties and types do not coexist in the same category, and a functor cannot be taken to be its own lifting. There is no choice but to construct a lifting. Details about fibrations can be found in standard references such as [8, 14].

Working in the general fibrational setting also facilitates a direct comparison of our work with that of Hermida and Jacobs [6]. The main difference is that they use fibred products and coproducts in defining their liftings, whereas we use left adjoints to reindexing functors instead. The codomain fibration over \( \text{Set} \) has both, so their derivation gives exactly the same induction rule as ours in the setting of Section 3.

Let \( U : \mathcal{E} \to \mathcal{B} \) be a fibration. Objects of the total category \( \mathcal{E} \) can be thought of as properties, objects of the base category \( \mathcal{B} \) can be thought of as types, and \( U \) can be thought of as mapping each property \( E \) in \( \mathcal{E} \) to the type \( UE \) of which \( E \) is a property. One fibration \( U \) can capture many different properties of the same type, so \( U \) is not injective on objects. For any object \( B \) of \( \mathcal{B} \), we write \( \mathcal{E}_B \) for the fibre above \( B \), i.e., for the subcategory of \( \mathcal{E} \) consisting of objects \( E \) such that \( UE = B \) and morphisms \( k \) between objects of \( \mathcal{E}_B \) such that \( Uk = id_B \). Let \( f^!_E \) be the cartesian morphism determined by \( f \) and \( E \). Then \( f^!_E \) is unique up to isomorphism for every choice of object \( E \) and morphism \( f \) with codomain \( UE \).

If \( f : B \to B' \) is a morphism in \( \mathcal{B} \), then the reindexing functor induced by \( f \) is the functor \( f^* : \mathcal{E}_{B'} \to \mathcal{E}_B \) defined on objects by \( f^* E = \text{dom}(f^!_E) \) and, for a morphism \( k : E \to E' \) in \( \mathcal{E}_{B'} \), \( f^*k \) is the morphism such that \( k \circ f^!_E = f^!_{E'} \circ f^*k \). The universal property of \( f^!_E \) ensures the existence and uniqueness of \( f^*k \).
Proceeding by analogy with the situation for Set-based predicates — where Set plays the role of B and P plays the role of E — we now define, for every functor F on B, a lifting ˆF of F to E such that U ˆF = FU. We construct ˆF by generalising each aspect of the construction of Section 3 to the general setting.

- **The Truth Functor** To construct the truth functor in the general setting, we assume B and E have terminal objects 1_B and 1_E, respectively, such that U(1_E) = 1_B. Writing !_B for the unique map from an object B of B to 1_B, we define K_1 : B → E by setting K_1B = (!_B)*1_E and, for a morphism k : B → B’, taking K_1k to be the unique morphism guaranteed to exist by the universal property of k!. Then, for every B in B, K_1B is the terminal object of E_B, so U(K_1B) = B. In fact, U ⊕ K_1, and the unit of this adjunction is id, so UK_1 = id.

- **Comprehension** Recall from Lemma 2 that the comprehension functor of Section 3 is right adjoint to the truth functor. Since right adjoints are defined up to isomorphism, in the general fibrational setting we can define the comprehension functor {−} to be the right adjoint to the truth functor K_1. A fibration U : E → B which has a right adjoint K_1 which itself has a right adjoint {−} is called a comprehension category [8]. We henceforth restrict attention to comprehension categories. We write ε for the counit of the adjunction {−} ⊕ K_1 and so, for any object E of E, we have that ε_E : K_1E → E.

- **Projection** Recall from Section 3 that the first step of the construction of our lifting is to define the projection π on this adjoint, which we denote Fπ. Putting this all together, we now define the lifting ˆF : E → E by ˆF_E = Σ_{FπE}(K_1(F\{E\})) for every object E of E. For completeness we also give the action of ˆF on morphisms. For each k : E → E’, define ˆFk = (FUk)^b ◦ K_1F(E’)^γ(K_1F\{E\}). Here, i) ◦ K_1F(E’) : Σ_{E_E}(F\{k\})^∗ K_1F(E’) → (FU\{k\})^∗ Σ_{E_E}(F\{k\})^∗ K_1F(E’), and ii) γ : Σ_{E_E}K_1F\{E\} → Σ_{E_E}(F\{k\})^∗ K_1F(E’) is the isomorphism arising from the fact that (F\{k\})^∗ is a right adjoint by the existence of Σ_{E_E} and hence preserves terminal objects. It is trivial to check that ˆF is indeed a lifting.
• Generalising Lemma 1 As in Section 3, we ultimately want to show that \( F \)-
algebras with carrier \( \{ P \} \) are interderivable with \( F \)-algebras with carrier \( P \). We first show that, as in Lemma 1, \( \hat{F}(K_1B) = K_1(FB) \) for any functor \( F \) on \( B \) and \( B \) in \( B \). Recall that \( UK_1 = id \) and define \( \pi K_1 : \{-\}K_1 \rightarrow Id \) to be the natural transformation with components \( \pi K_1 )^{-1} = UK_1^{-1} \), where \( \eta : Id \rightarrow \{-\}K_1 \) is the unit of the adjunction \( \{-\} \dashv K_1 \), so that \( \pi K_1 \) is a natural isomorphism. If we further define \( F\pi K_1 : F\{-\}K_1 \rightarrow FK_1 \) to be the natural transformation with components \( (F\pi K_1)_B = F((\pi K_1)_B) \), then \( F\pi K_1 \) is also a natural isomorphism. We will use this observation below to show that, for every object \( B \) of \( B \), \( \Sigma_{(F\pi K_1)_B} \) is not only left adjoint by definition, but also right adjoint, to \( (F\pi K_1)_B^\ast \). Then, observing that right adjoints preserve terminal objects and that \( K_1(FB) \) is the terminal object of \( E_{FB} \) (since \( K_1B \) is the terminal object of \( E_B \) for any \( B \)), we will have shown that \( \hat{F}(K_1B) = \) i.e., \( \Sigma_{(F\pi K_1)_B}(K_1(F\{K_1B\})) \) — must be the terminal object of \( E_{FB(K_1B)} \), i.e., of \( E_{FB} \). In other words, we will have shown that \( \hat{F}(K_1B) = K_1(FB) \).

So, fix an object \( B \) of \( B \). To see that \( \Sigma_{(F\pi K_1)_B} \vdash ((F\pi K_1)_B)^\ast \), first note that, for any isomorphism \( f : B \rightarrow B' \) in \( B \), \( f^\ast \) and \( (f^{-1})^\ast \) both exist and both \( f^\ast \vdash (f^{-1})^\ast \) and \( (f^{-1})^\ast \vdash f^\ast \) hold. Then, since \( f^\ast \vdash \Sigma_f \) by definition, we have \( \Sigma_f = (f^{-1})^\ast \), and thus that \( \Sigma_f \vdash f^\ast \). Instantiating \( f \) to \( (F\pi K_1)_B \) and recalling that \( (F\pi K_1)_B \) is an isomorphism, we have that \( \Sigma_{(F\pi K_1)_B} \vdash ((F\pi K_1)_B)^\ast \).

• A Generic Fibrational Induction Rule Analogues of Lemma 3 and Lemma 4 hold in the general fibrational setting provided all occurrences of \( \text{Set} \) are replaced by \( B \) and all occurrences of \( \mathcal{P} \) are replaced by \( E \) and, in the analogue of Lemma 2, \( \{-\} : B(B, \{E\}) \rightarrow E(K_1B, E) \) and \( \{-\}^\# : E(K_1B, E) \rightarrow B(B, \{E\}) \) are defined by the adjunction \( \{-\} \vdash K_1 \).

The above construction thus yields the following generalised induction rule:

\[
\text{genfibind} \quad F \in \mathcal{B} \quad (E : E_X) \quad (k : \hat{F}E \rightarrow E) \quad \mu F \rightarrow \{E\}
\]

\[
\text{genfibind} \quad F \quad E \quad = \quad \text{fold} \circ \Psi
\]

This induction rule looks slightly different from the one for set-valued predicates. In Section 3, we were able to use the specific structure of comprehensions for set-valued predicates to extract proofs for particular data elements from them. But in the general fibrational setting, predicates, and hence comprehensions, are left abstract, so we take the return type of the general induction scheme \( \text{genfibind} \) to be a comprehension. We predict that, when \( \text{genfibind} \) is instantiated to a fibration of interest, we should be able to use knowledge about that fibration to extract from the comprehension it constructs further proof information relevant to the application at hand. This expectation is justified, as in [6], by \( \{-\} \vdash K_1 \).

We now give an induction rule for a data type and properties that cannot be modelled in \( \text{Set} \).

**Example 4** The fixed point \( H_{\text{yp}} = \mu F \) of the functor \( F = (X \rightarrow \text{Int}) \rightarrow \text{Int} \) is the data type of hyperfunctions. Since \( F \) has no fixed point in \( \text{Set} \), we interpret it in the category \( \omega CPO_\perp \) of \( \omega \)-cpos with \( \perp \) and strict continuous monotone functions. In this setting, a property of an object \( X \) of \( \omega CPO_\perp \) is an admissible
sub-$\omega CPO$ $A$ of $X$. Admissibility means that the bottom element of $X$ is in $A$ and $A$ is closed under the least upper bounds of $X$. This structure forms a Lawvere category [7, 8]. The truth functor maps $X$ to $X$, and comprehension maps a sub-$\omega CPO$ $A$ of $X$ to $A$. The lifting $\hat{F}$ maps a sub-$\omega CPO$ $A$ of $X$ to the sub-$\omega CPO$ $FA$ of $FX$. Finally, the derived induction rule states that if $A$ is an admissible sub-$\omega CPO$ of $\text{Hyp}$, and if $\hat{F}(A) \subseteq A$, then $A = \text{Hyp}$.

5 Conclusion and future work

We give an induction rule that can be used to prove properties of data structures of inductive types. Like Hermida and Jacobs, we give a fibrational account of induction, but we derive, under slightly different assumptions on fibrations, a generic induction rule that can be instantiated to any inductive type rather than just to polynomial ones. This rule is based on initial algebra semantics of data types, and is parameterised over both the data types and the properties involved. It is also principled, expressive, and correct. Our derivation yields the same induction rules as Hermida and Jacobs’ when specialised to polynomial functors in the codomain fibration, but it also gives induction rules for non-polynomial data types such as rose trees, and for data types such as finite hereditary sets and hyperfunctions, for which no induction rules have previously been known.

There are several directions for future work. The most immediate is to instantiate our theory to give induction rules for nested types. These are exemplified by the data type of perfect trees given in Haskell-like syntax as follows:

```haskell
data PTree a : Set where
  PLeaf : a → PTree a
  PNode : PTree (a, a) → PTree a
```

Nested types arise as least fixed points of rank-2 functors; for example, the type of perfect trees is $\mu H$ for the functor $H$ given by $HF = \lambda X. X + F(X \times X)$. An appropriate fibration for induction rules for nested types thus takes $B$ to be the category of functors on $\text{Set}$, $E$ to be the category of functors from $\text{Set}$ to $\mathcal{P}$, and $U$ to be postcomposition with the forgetful functor from Section 3. A lifting $H$ of $H$ is given by $H PX (\text{inl } a) = 1$ and $H PX (\text{inr } n) = P (X \times X) n$. Taking the premise to be an $\hat{H}$-algebra gives the following induction rule for perfect trees:

$$\text{indPTree} : \forall (P : \text{Set} → \mathcal{P}).
\forall (X : \text{Set})(x : X). P(\text{PLeaf } x) →
\forall (X : \text{Set})(t : \text{PTree}(X \times X)). P(\text{X} t → P(\text{PNode } t)) →
\forall (X : \text{Set})(t : \text{PTree } X). P X t$$

This rule can be used to show, for example, that $\text{PTree}$ is a functor.

Extending the above instantiation for the codomain fibration to “truly nested types” and fibrations is current work. We expect to be able to instantiate our theory for truly nested types, GADTs, indexed containers, and dependent types, but initial investigations show care is needed. We must ascertain which fibrations can model predicates on such types, since the codomain fibration may not
give useful induction rules, as well as how to translate the rules to which these fibrations give rise to an intensional setting.

Matthes [11] gives induction rules for nested types (including truly nested ones) in an intensional type theory. He handles only rank-2 functors that underlie nested types (while we handle any functor of any rank with an initial algebra), but his insights may help guide choices of fibrations for truly nested types. These may in turn inform choices for GADTs, indexed containers, and dependent types.

Induction rules can automatically be generated in many type theories. Within the Calculus of Constructions [3] an induction rule for a data type can be generated solely from the inductive structure of that type. Such generation is also a key idea in the Coq proof assistant [4]. But as far as we know, generation can currently be done only for syntactic classes of functors rather than for all functors with initial algebras. In some type theories induction schemes are added as axioms rather than generated. For example, attempts to generate induction schemes based on Church encodings in the Calculus of Constructions proved unsuccessful and so initiality was added to the system, thus giving the Calculus of Inductive Constructions. Whereas Matthes’ work is based on concepts such as impredicativity and induction recursion rather than initial algebras, ours reduces induction to initiality, and may therefore help lay the groundwork for extending implementations of induction to more sophisticated data types.

References

4. The Coq Proof Assistant. Available at coq.inria.fr